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Abstract

This report summarizes and consolidates the main results from OpenTC Workpackage 05 “Security Management and Infrastructure”. The goal of Workpackage 5 has been to develop mechanisms for managing security of virtual systems while leveraging trusted computing technologies for verifiability and protection.

Part I of the deliverable introduces our goals and surveys the approach we have taken from a high-level perspective. Part II then describes key building blocks in detail. Part III describes the WP5 concepts in the demonstrator and concludes the deliverable by evaluating our prototype and documenting our future outlook.

Contributors and Acknowledgements

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Contents

Chapter 1

Introduction and Outline

1.1 Introduction

Hardware virtualization is enjoying a resurgence of interest fueled in part by its cost-saving potential. By allowing multiple virtual machines to be hosted on a single physical server, virtualization helps improve server utilization, reduce management and power costs, and control the problem of server sprawl.

A prominent example in this context is data centers. The *infrastructure provider*, who owns, runs, and manages the data center, can transfer the cost savings to its customers or *outsourcing companies*, whose virtual infrastructures are hosted on the data center's physical resources. A large number of the companies that outsource their operations are small and medium businesses or SMBs, which cannot afford the costs of a dedicated data center in which all the data center's resources are used to host a single company's IT infrastructure. Hence, the IT infrastructure belonging to multiple SMBs may be hosted inside the same data center facility. Today, even in such "shared" data centers, each run on distinct physical resources and there is no resource sharing among various customers. In this so-called *physical cages* model, the customers are physically isolated from each other in the same data center.

Limited trust in the security of virtual datacenters is one major reason for customers not sharing physical resources. Since management is usually performed manually, administrative errors are commonplace. While this may lead to down times in virtual datacenters used by a single customer, it can lead to information leakages to competitors if the datacenter is shared. Furthermore, multiple organizations will only allow sharing of physical resources if they can trust that security incidents cannot spread across the isolation boundary separating two customers.

Security Objectives Our main security objective is to provide isolation among different domains that is comparable[1] with the isolation obtained by providing one infrastructure for each customer. In particular, we require a security architecture that protects those system components that provide the required isolation or allow to verifiably reason about their trustworthiness of and also of any peer endpoint (local or remote) with a domain, i.e., whether they conforms to the underlying security policy.

We achieve this by grouping VMs dispersed across multiple physical resources into a *virtual zone* in which customer-specified security requirements are automatically

[1] Note that unlike physical isolation, we do not solve the problem of covert channels.

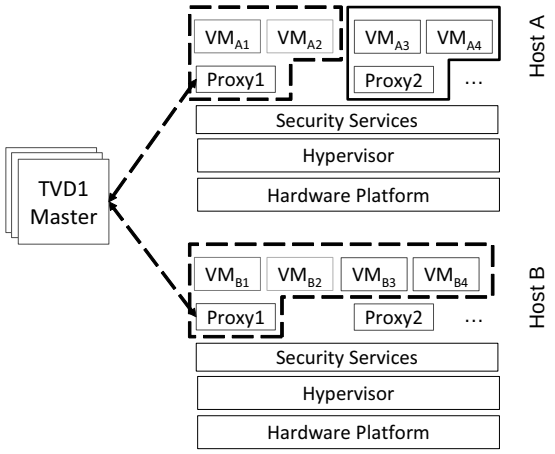

Figure 1.1: TVD Architecture: High-Level Overview.

enforced. Even if VMs are migrated (say, for load-balancing purposes) the logical topology reflected by the virtual domain should remain unchanged. We deploy Trusted Computing (TC) functionalities to determine the trustworthiness (assure the integrity) of the policy enforcement components.

Such a model would provide better flexibility, adaptability, cost savings than today's physical cages model while still providing the main security guarantees required for applications such as datacenters.

1.2 Our Contribution

In this deliverable, we provide a blueprint for realizing integrity and isolation in virtual systems. We do this by supporting a logical cages model, in particular for virtualized data centers, based on a concept called Trusted Virtual Domains or TVDs [16]. Based on previous work, we describe a security management framework that helps to realize the abstraction of TVDs by guaranteeing reliable isolation and flow control between domain boundaries. Our framework employs networking and storage virtualization technologies as well as Trusted Computing for policy verification. Our main contributions are (1) combining these technologies to realize TVDs and (2) orchestrating them through a management framework that automatically enforces isolation among different zones. In particular, our solution aims at automating the verification, instantiation and deployment of the appropriate security mechanisms and virtualization technologies based on an input security model, which specifies the required level of isolation and permitted information flows.

1.3 Outline of this Report

This report is structured in three parts. Part I surveys our work and summarizes related work in Chapter 4. Part II describes selected building blocks of our security architecture in detail. Part III describes WP5 components of our prototype and finally reflects on

our lessons learned and documents our outlook onto the future of trusted computing and virtualization.

The first technical component in Part II is the integrity and assurance management of the OpenTC Security Services. This has two aspects: In Section 5 we describe how integrity statements about virtual machines can be made and how data can be bound to the integrity of a machine. We also describe how to protect the privacy of users using our system. In Section 6 we extend these results to cover hierarchical integrity management, i.e., the integrity protection of packages of multiple virtual machines and the related components.

The second component is our network security architecture described in Chapter 7. It implements two key ideas. The first idea is to provide secure virtual networks (so-called trusted virtual domains) that transparently connect VMs on multiple hosts. The second idea is to automatically provision the required protection mechanisms such that the networks guarantee a given set of user requirements.

The third component described in Chapter 8 is the public key infrastructure (PKI) that manages the keys used four security mechanisms. It also includes trusted computing extensions to existing PKI standards.

The fourth component are trusted channels in Chapter 9 that enable to establish a secure channel while verifying the integrity of the peer. This allows users to not only guarantee the integrity of a given machine but also to securely connect to the machine that has been validated.

The fifth and final concept is our approach to attack and failure resilience as document in Chapter 10. This concept comprises three main ideas. The first is to use introspection to detect viruses and other failures inside a running VM. The second is to implement redundancy for VMs and key hypervisor components such as network or disk drivers. The final idea combining them is to monitor VMs and rejuvenate VMs that are failed or are at risk of failing. Overall, this allows a substantial increase in the resiliency of the services running on this platform.

Part I

Overview and Related Work

1.4 Dependability of Virtual Systems

We now provide a sampling of related work in the area of using VMs for improving dependability. We also compare our X-Spy intrusion detection framework with previous hypervisor-based intrusion detection systems. Many of these works, including ours, implicitly trust the virtualization layer to function properly, to isolate the VMs from each other, and to control the privileged access of certain VMs to other VMs. Such a trust can be justified by the observation that a typical hypervisor consists of some tens of thousands lines-of-code (LOC), whereas a typical operating system today is on the order of millions LOC [40]. This allows a much higher assurance for the code of a hypervisor.

Bressoud and Schneider [15] implemented a primary-backup replication protocol tolerant to benign faults at the VMM level. The protocol resolves non-determinism by logging the results of all non-deterministic actions taken by the primary and then applying the same results at the backups to maintain state consistency.

Double-Take [124] uses hardware-based real-time synchronous replication to replicate application data from multiple VMs to a single physical machine so that the application can automatically fail over to a spare machine by importing the replicated data in case of an outage. As the replication is done at the file system level below the VM, the technique is guest-OS-agnostic. Such a design could provide the basis for a business model in which multiple client companies outsource their disaster recovery capability to a disaster recovery hot-site that houses multiple physical backup machines, one for each client.

Douceur and Howell [29] describe how VMMs can be used to ensure that VMs satisfy determinism and thereby enable state machine replication at the VM level rather than the application level. Specifically, they describe how a VM's virtual disk and clock can be made deterministic with respect to the VM's execution. The design relieves the application programmer of the burden of structuring the application as a deterministic state machine. Their work is similar to Bressoud and Schneider's approach [15] of using a VMM to resolve non-determinism. However, the difference lies in the fact that whereas Bressoud and Schneider's approach resolves non-determinism using the results of the primary machine's computation, Douceur and Howell's design resolves non-determinism *a priori* by constraining the behavior of the computation.

Dunlap *et al.* describe ReVirt [31] for VM logging and replay. ReVirt encapsulates the OS as a VM, logs non-deterministic events that affect the VM's execution, and uses the logged data to replay the VM's execution later. Such a capability is useful to recreate the effects of non-deterministic attacks, as they show later in [59]. Their replay technique is to start from a checkpoint state and then roll forward using the log to reach the desired state.

Joshi *et al.* [59] combine VM introspection with VM replay to analyze whether a vulnerability was activated in a VM before a patch was applied. The analysis is based on vulnerability-specific predicates provided by the patch writer. After the patch has been applied, the same predicates can be used during the VM's normal execution to detect and respond to attacks.

Backtracker [64] can be used to identify which application running inside a VM was exploited on a given host. Backtracker consists of an online component that records OS objects (such as processes and files) and events (such as read, write, and fork), and an offline component that generates graphs depicting the possible chain of events between the point at which the exploit occurred and the point at which the exploit was detected.

An extension of Backtracker [66] has been used to track attacks from a single host

at which an infection has been detected to the originator of the attack and to other hosts that were compromised from that host. The extension is based on identifying causal relationships, and has also been used for correlating alerts from multiple intrusion detection systems.

King *et al.* [65] describe the concept of time-traveling virtual machines (TTVMs), in which VM replay is used for low-overhead reverse debugging of operating systems and for providing debugging operations such as reverse break point, reverse watch point, and reverse single step. Combining efficient checkpointing techniques with ReVirt, TTVMs can be used by programmers to go to a particular point in the execution history of a given run of the OS. To recreate all relevant state for that point, TTVMs log all sources of non-determinism.

Garfinkel and Rosenblum [40] introduced the idea of hypervisor-based intrusion detection, and pointed out the advantages of this approach and its applicability not only for detection, but also for protection. Their Livewire system uses a modified VMware workstation as hypervisor and implements various polling-based and event-driven sensors. Compared with Livewire, our X-Spy system employs more extensive detection techniques (e.g., by checking not only processes, but also kernel modules and file systems) and protection techniques (such as pre-checking and white-listing of binaries, and kernel sealing) with an explicit focus on rootkit detection. In addition, X-Spy enables easy forensic analysis.

Zhang *et al.* [132] and Petroni *et al.* [85] use a secure coprocessor as the basis for checking the integrity of the OS kernel running on the main processor. However, as the coprocessor can only read the memory of the machine monitored, only polling-based intrusion detection is possible. In contrast, X-Spy can perform both polling-based and event-driven intrusion detection. Specifically, it can intercept and deny certain requested actions (such as suspicious system calls), and therefore has the capability to not only detect but also protect.

Laureano *et al.* [71] employ behavior-based detection of anomalous system call sequences after a learning phase in which "normal" system calls are identified. Processes with anomalous system call sequences are labeled suspicious. For these processes, certain dangerous system calls will in turn be blocked. The authors describe a prototype based on a type-II hypervisor, namely, User-Mode Linux (UML) [28].

The ISIS system of Litty [75] is also based on UML. ISIS runs as a process in the host operating system and detects intrusions in the guest operating system by using the `ptrace` system call for instrumenting the guest UML kernel. Unlike X-Spy, ISIS focuses mostly on intrusion detection and not protection.

Jiang *et al.* [57] describe the *VMwatcher* system, in which host-based anti-malware software is used to monitor a VM from within a different VM. X-Spy and VMwatcher are similar in that both use the hypervisor as a bridge for cross-VM inspection, and both tackle the semantic gap problem. While their work focuses on bridging the semantic gap on a multitude of platforms (hypervisors and operating systems), our work focuses on employing more extensive detection mechanisms (such as checking not only processes, but also kernel modules, network connections, and file systems) on a single hypervisor. In contrast to X-Spy, VMwatcher does not include event-driven detection methods or protection techniques.

The Strider GhostBuster system by Beck *et al.* [12] is similar to X-Spy in that both use a differential view of system resources. Strider GhostBuster compares high-level information (such as information obtained by an OS command) with low-level information (e.g., kernel information) to detect malicious software trying to hide system resources from the user and administrator. However, such a comparison has limited

effectiveness as detection takes place in the same (potentially compromised) OS environment. Beck *et al.* also compare the file system view obtained from a potentially compromised OS with the view obtained from an OS booted from a clean media. The disadvantage of such an approach is that it requires multiple reboots and is limited to checking only persistent data (such as file system) and not run-time data.

Chapter 2

Security Policies for Virtual Data Centers

Data centers provide computing and storage services to multiple customers. Customers are ideally given dedicated resources such as storage and physical machines. In the physical cages approach, only few resources such as the Internet connection may be shared between multiple customers. For cost efficiency, our logical cages approach promotes securely extending sharing to other resources such as storage and networks. This is enabled by preventing unauthorized information exchange via shared resources.

To model and implement the logical caging approach, we introduce a domain-based security model for enforcing unified security policies in virtualized data centers. We focus on isolation policies that mimic physical separation of data center customers. Our goal is to logically separate networks, storage, VMs, users, and other virtual devices of one customer from another customer. For our purposes, we define *domain isolation* as the ability to enforce security policies within a domain independently of other domains that may co-exist on the same infrastructure and interact with that domain. The core idea is to use this isolation property as a foundation for guaranteeing desired security properties within each virtual domain while managing shared services under mutually agreed policies.

We now explain the policies that describe this controlled information exchange in a virtualized data center. In Section 3 we describe the individual components that enable us to enforce these policies.

2.1 High-level Policy Model

The security model includes two high-level policies defining the security objectives that must be provided by the underlying infrastructure:

Inter-TVD Policy: By default, each TVD is isolated from the outside world. The high-level information-exchange policy defines whether and how information can be exchanged with other TVDs. If no information flow with other TVDs is permitted, no resources can be shared unless the data center operator can guarantee that the isolation is preserved. If information flow to/from other TVDs is allowed, sub-policies further qualify the exact information flow policy for the individual resources.

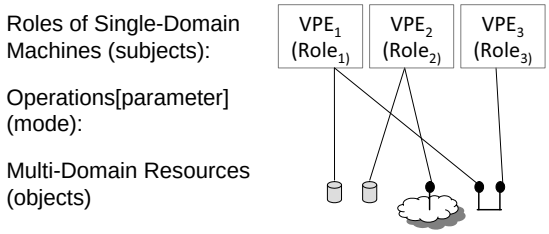

Figure 2.1: Policy Model: Single-TVD Machines operate on Shared Resources

Intra-TVD Policy: Domain policies allow TVD owners (e.g., customers) to define the security objectives within their own TVDs. Examples of such policies include how the internal communication is to be protected and under what conditions resources (e.g., storage, machines) can join a particular TVD.

We further define more fine-grained policies by the use of *roles* that can be assigned to any member VM, say to a member machine. This allows us to define and enforce role-based policies within and across TVDs. For example, machines can now assume internal or gateway roles with corresponding permissions; while a workstation may not be allowed to connect to non-TVD networks, machines with the "firewall" role can be allowed to connect to selected other networks. Figure 2.1 depicts three VMs in a single TVD. Each VM is given different levels of access to resources with respect to their role for that TVD.

2.2 Security Objectives and Policy Enforcement Points

Policies are enforced for all shared resources in the TVD infrastructure (see Figure 2.2). The basis of all policies is isolation at the boundary of each TVD. By default, each resource is associated with a single domain. This achieves a basic level of isolation. If information flow between TVDs is allowed, resources can also be member of different TVDs. For example, a TVD can allow certain types of resources on certain hosts to provide services also to other domains. Each TVD defines rules regarding in-bound and out-bound information flow for restricting communication with the outside world. The underlying policy-enforcement infrastructure then has to ensure that only resources trusted by all TVDs are shared.

Architecturally, there are two ways of enforcing such rules, depending on the trust between the TVDs. The first method involves two shared resources connected by an intermediate domain. In this method, each TVD enforces its side of the flow control by means of its own shared resource. An example of this type of connection is the one that exists between *TVD A* and *TVD B* in Figure 2.2. This method is used when the trust level between *TVD A* and *TVD B* is low, and the two cannot agree on a shared resource that is mutually trusted. The shared resource in *TVD A* will enforce *TVD A*'s policies regarding in-bound traffic from *TVD B*, even if the shared resource in *TVD B* does not enforce *TVD B*'s policies regarding out-bound traffic. The shared resources can be thought of being a part of a "neutral" TVD (*TVD AB*) with its own set of membership requirements. The second method that requires shared trust is to establish one or more shared resources that are accessed from both TVDs while allowing controlled information flow. This mechanism is used between *TVD B* and *TVD C* in Figure 2.2.

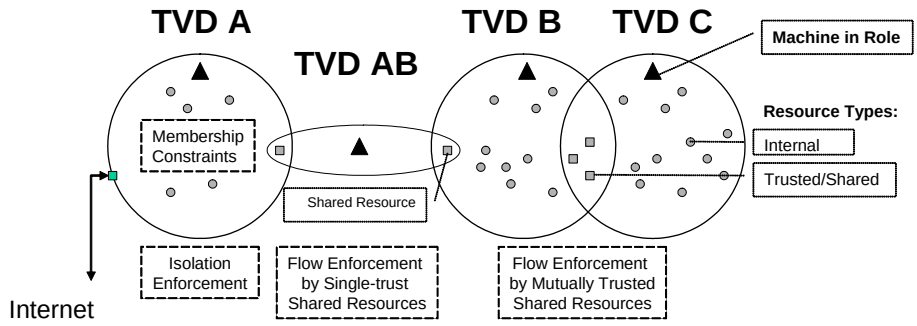

Figure 2.2: Usage Control for Shared Resources: Machines use resources belonging to TVDs.

From / to	D_I	D_D	D_i
D_I	**1**	1	0
D_D	0	**1**	1
D_i	0	1	**1**

Table 2.1: High-level Directed Flow Control Matrix for Internet D_I, DMZ D_D, and Intranet D_i.

Security within a virtual domain is finally obtained by defining and enforcing *membership requirements* that resources have to satisfy prior to being admitted to the TVD and for retaining the membership. This may also include special requirements for different machine types: Because, for example, shared resources play a key role in restricting information flow between TVDs, the software on those machines may be subject to additional integrity verification as compared to the software on regular VMs.

2.2.1 Permitted Flows in Data Centers

At a high level flow control policies define the allowed traffic flow between two domains and how the domains should be protected. Allowed information flows can be represented by a simple flow control matrix as depicted in Table 2.1, where 1 allows information flow and 0 denies it. This example implements a basic enterprise policy that regulates incoming flow from untrusted outside entities (D_I) through a semi-trusted intermediary domain (D_D), and disallows any outgoing flow. Note that this matrix is directional, i.e., it might allow flows in one direction but not in the opposite direction. If flow policies between two TVDs are asymmetric, only shared resources that can enforce these policies are permitted.

Device-specific policies (network, storage) can then refine these basic rules. If an information flow is not permitted, then also shared resources are not permitted between these TVDs.

2.2.2 Membership Requirements

Membership requirements define under what conditions resources may join a domain. From a high-level policy perspective, several criteria can be applied to decide whether an entity is allowed to join a domain, for example:

- *Certificates*: An authority defined by the TVD policy can certify a resource to be member of a TVD. A common example is that an enterprise issues machine certificates to allow its machines to join the corporate network.
- *Integrity Proofs*: A resource may prove its right to join a TVD using integrity proofs. It may, e.g., prove that the integrity of the base operating system is intact and that all required patches have been applied [103].
- *User-identity*: Only machines operated by a certain user can join. This can be validated by user-name/password or by a cryptographic token.

In general, a resource may need to show proper credentials to prove that it fulfills certain properties before allowing the resource to join the TVD [96]. More formally, a machine m is permitted to join a TVD t if and only if there is at least one property of m that satisfies each security requirement of t. The validations of these properties are usually done on a per-type and role basis. For example, requirements for a shared resource are usually stronger than the requirements for a TVD-internal resource.

2.3 Example Policy Refinements for Protected Resources

Policies alone are not sufficient to enforce customer separation in a virtualized data center. Ultimately, one needs to transform these policies into data center configurations and security mechanisms specific to each resource (e.g., VLAN configuration). To do so, we introduce a policy management scheme that accepts high-level domain policies and transforms them into resource-specific low-level policies and configurations. In Section 11 we demonstrate a prototype based on this architecture that enforces high-level TVD policies by lower-level network and infrastructure configurations, which is then deployed onto each physical platform to assist customer separation.

2.3.1 Refinement Model

The high-level policy defines the basic flow control, protection, and admission requirements. We aim at enforcing these high-level objectives throughout all resources in the data center.

In the high-level model, flow control across customer domains is specified by a simple matrix such as the one in Figure 2.1 that defines whether flows are permitted. This however is not sufficiently fine-grained for specific resources. TVDs, for example, want to restrict their flow across boundaries by means of firewall rules. As a consequence, we need to introduce a notion of policy refinement [127], because as translation moves towards lower levels of abstraction, it will require additional information (e.g., physical arrangement of the data center, “subjective” trust information) to be correctly and coherently executed.

Our notion of policy refinement mandates the enforcement of “no flow” objectives while allowing each resource to refine what it means so that flows are permitted and how exactly unauthorized flows shall be prevented. Similarly, we do not allow resources to deviate from the confidentiality/integrity objectives; however, certain resources can be declared trusted so that they may enforce these objectives without additional security mechanisms such as encryption or authentication.

Flow to $\rightarrow$	D_I		D_D		D_i	
Enforced by $\downarrow$	gate	internal	gate	internal	gate	internal
D_I	**1**	**1**	P_{ID}	0	**0**	**0**
D_D	0	0	**1**	**1**	P_{Di}	0
D_i	**0**	**0**	P_{Di}	0	**1**	**1**

Table 2.2: Example Network Flow Control Policy Matrix for Three TVDs.

Similarly, the fact that admission is restricted is then refined by specific admission control policies that are enforced by the underlying infrastructure.

Note that conflict detection and resolution [127, 76] can later be used to extend this simple notion of refinement. However, we currently stay on the safe side: Connections are only possible if both TVDs allow them. Similarly, if one domain requires confidentiality, information flows are only allowed to TVDs that also require confidentiality. Other schemes for more elaborate flow control have been proposed in [33, 17, 32, 38].

2.3.2 Network Security Policies

We now survey the policy model of [18] and show how it related to the corresponding high-level policy. Similar to our high-level policies, there are two types of policies governing security in the network. The first limits flow between networks, whereas the second defines membership requirements to each network.

Network Security Policies across TVDs A policy covers isolation and flow control between TVDs as well as integrity and confidentiality against outsiders. These basic security requirements are then mapped to appropriate policies for each resource. For example, from a networking perspective, isolation refers to the requirement that, unless the inter-TVD policies explicitly allow such an information flow, a dishonest VM in one TVD cannot (1) send messages to a dishonest VM in another TVD (information flow), (2) read messages sent on another TVD (confidentiality), (3) alter messages transmitted on another TVD (data integrity), and (4) become a member of another TVD network (access control).

TVDs often constitute independent organizational units that may not trust each other. If this is the case, a communication using another TVD can be established (see the communication between TVD A and B in Figure 2.2). The advantage of such a decentralized enforcement approach is that each TVD is shielded from security failures in other TVDs, thus contribute to domain isolation. For networks, the main inter-TVD security objectives are controlled information sharing among the TVDs as well as integrity and confidentiality protection of the channel.

While the high-level model specifies whether information exchange is allowed between domains or not, we now refine this policy as follows:

- We refine the active elements (subjects) of given domains by introducing roles that machines can play. This allows us to set different permissions to boundary machines as compared to internal machines.
- In case information flow is permitted in principle, we refine the network security policies by introducing flow control rules that can further restrict the actual

information exchange. A network policy may disallow flow even though it has been allowed from a high-level policy perspective.

An information flow control matrix is a simple way of formalizing these network connectivity objectives. Table 2.2 shows a sample matrix for the three example TVDs introduced earlier. Each matrix element represents a policy specifying permitted connections between a pair of TVDs, as enforced by one of the TVDs. The depicted policies P_x that limit information exchange will be implemented by firewall rules that are used to program the boundary firewalls. The **1** values along the matrix diagonal convey the fact that there is free information exchange within each TVD. The **0** values in the matrix are used to specify that there should be no direct information flow between two TVDs, e.g., between the Internet D_I and the intranet D_i. Care must be taken to ensure that the pairwise TVD policies specified in the information flow control matrix do not accidentally contradict each other or allow undesired indirect flow.

Intra-TVD Network Security Policy Within a TVD, all VMs can freely communicate with each other while observing TVD-specific integrity and confidentiality requirements. For this purpose, the underlying infrastructure may ensure that intra-TVD communication only takes place over an authenticated and encrypted channel (e.g., IPSec), or alternatively, a trusted network[1].

2.3.3 Towards Storage Security Policies

Virtual disks attached to VMs must retain the advantages offered by storage virtualization while at the same time enforcing TVD security policies. Advantages of storage virtualization include improved storage utilization, simplified storage administration, and the flexibility to accommodate heterogeneous physical storage devices. Similar to network, we now show a refinement of the high-level TVD policies into access control policies for VMs in certain roles to disks belonging to a domain.

Inter-TVD Storage Security A virtual disk has a single label corresponding to the TVD it belongs to. Whenever a virtual machine operates on virtual storage, the global flow matrix described in Section 2 needs to be satisfied. For flexibility, each TVD can define a set of storage policies that govern usage and security of its storage. A single policy is then assigned to and enforced for each storage volume.

As the starting point of our storage policy refinement, we define a *maximum permission policy* as follows:

1. Any machine in domain TVD_A playing any role can write to a disk of domain TVD_B iff flow from domain TVD_A to domain TVD_B is permitted.

2. Any machine in domain TVD_A playing any role can read from a disk of domain TVD_B iff flow from domain TVD_B to domain TVD_A is permitted.

3. Any single machine in any domain can read/write mount a blank disk. After data is written, the disk changes ownership and is now assigned to the domain of the machine who has written data.

[1] A network is called *trusted* with respect to a TVD security objective if it is trusted to enforce the given objective transparently. For example, a server-internal Ethernet can often be assumed to provide confidentiality without any need for encryption.

Flow to $\rightarrow$	D_I		D_D		D_i	
Disk $\downarrow$	gate	internal	gate	internal	gate	internal
D_I	r/w	r/w	w	0	**0**	**0**
D_D	r	0	r/w	r/w	r/w	0
D_i	**0**	**0**	r/w	0	r/w	r/w
Blank	$r/$	0	$r/$	0	$r/$	0
	$w \rightarrow D_I$	0	$w \rightarrow D_D$	0	$w \rightarrow D_i$	0

Table 2.3: Example of a Refined Disk Policy Matrix for Three TVDs.

Table 2.3 shows the resulting maximum disk access control policy. Actual policies are then valid with respect to a maximum-permission policy for a domain if they permit a subset of its permissions. Note that as flow within a domain is always allowed, this implies that disks of the same domain as the machine may always be mounted read/write.

Intra-TVD Storage Security By default, we consider the content of a disk to be confidential while the storage medium (possibly remote) is deemed to be untrusted. As a consequence, if a given domain does not declare a given storage medium as trusted, we deploy whole-disk encryption using a key that is maintained by the TVD infrastructure.[2] Another aspect reflected in the disk policies is the fact that we have a notion of blank disks. Once they are written by another domain, they change color, and are then associated with this other domain while being encrypted under the corresponding key. In the future, it would be desirable to have integrity-protected storage [24, 89] where the TVD can validate that its content have not been changed by untrusted entities.

For protecting the data in a particular TVD, virtual storage may in addition specify which conditions on the system must be satisfied before a disk may be *re-mounted* by a VM that has previously unmounted the disk, and whether shared mounting by multiple systems is allowed. Note that these membership restrictions require bookkeeping of disks and management of access of VMs to disks.

[2]Note that the VM only sees unencrypted storage, i.e., the TVD infrastructure automatically loops in encryption.

Chapter 3

Unified Policy Enforcement for Virtual Data Centers

In this section, we introduce a TVD-based policy enforcement framework that orchestrates the deployment and enforcement of the type of policies we presented in Section 2 across the data center. Existing storage and network virtualization technologies as well as existing Trusted Computing components (in software and hardware) are the building blocks of our solution. Our framework (1) combines these technologies to realize TVDs and (2) orchestrates them using the TVD infrastructure, which provisions the appropriate security mechanisms.

3.1 TVD Infrastructure

The TVD infrastructure consists of a management layer and an enforcement layer. The TVD management layer includes TVD masters, proxies, and factories, whereas the TVD enforcement layer consists of various security services. Each TVD is identified by a unique *TVD Master* that orchestrates TVD deployment and configuration. The TVD Master can be implemented as a centralized entity (as in our prototype described in Section 11) or have a distributed fault-tolerant implementation. The TVD Master contains a repository of high-level TVD policies and credentials (e.g., VPN keys). The Master also exposes a TVD management API through which the TVD owner can specify those policies and credentials. In the deployment phase, the TVD Master first verifies the suitability and capability of the physical host (which we refer to as pre-admission control). It then uses a generic *TVD Factory* service to spawn a *TVD Proxy*, which acts as the local delegate of the TVD Master dedicated to that particular host. The TVD Proxy is responsible for (1) translation of high-level TVD policies into low-level platform-specific configurations, (2) configuration of the host and security services with respect to the translated policies, and (3) interaction with the security services in TVD admission and flow control.

Security services implement the security enforcement layer of our TVD infrastructure. They run in a trusted execution environment on each physical host (e.g., Domain-0 in Xen) and (1) manage the security configuration of the hypervisor, (2) provide secure virtualization of resources (e.g., virtual devices) to the VMs, and (3) provide support to TVD proxies in enforcing flow and access control policies within and across TVD boundaries. Figure 3.1 shows a high-level list of security services and their interac-

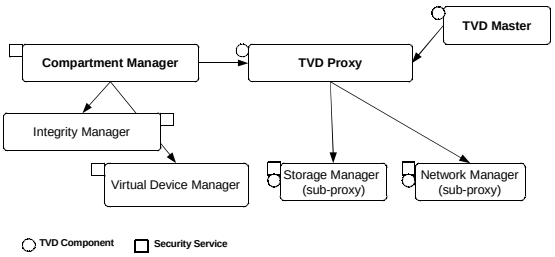

Figure 3.1: TVD Components and Security Services.

tion with the TVD components. Most importantly, the *compartment manager* service manages the life-cycle of VMs in both para-virtualized and fully virtualized modes. This service works in collaboration with the TVD Proxy to admit VMs into TVDs. The *integrity manager* service implements Trusted Computing extensions and assists the TVD Proxy in host pre-admission and VM admission control. The *virtual network manager* and *virtual storage manager* services are invoked by the TVD Proxy. They implement resource virtualization technologies and enforce parts of the high-level TVD policies that are relevant to their operation. Lastly, the *virtual device manager* service handles the secure resource allocation and setup of virtual devices assigned to each VM.

Our TVD infrastructure is geared towards automated deployment and enforcement of security policies specified by the TVD Master. Automated refinement and translation of high-level policies into low-level configurations are of particular interest. For example, for information flow between two hosts in a trusted data center environment, other mechanisms need to be in place than for a flow between two hosts at opposite ends of an untrusted WAN link. In the latter case, the hosts should be configured to allow communication between them only through a VPN tunnel.

Another important consideration is policy conflict detection and resolution [127, 76]. In fact, conflicting high-level policies (e.g., a connection being allowed in the inter-TVD policy but disallowed in the intra-TVD policy) can potentially result in an incorrect configuration of the underlying infrastructure. We cannot solely rely on the TVD owner to specify conflict-free policies. It is important to detect policy conflicts and provide feedback to the owner in case one is detected. In the present prototype, policy refinement is performed manually. The result is a set of configuration files that we use for configuring the security services at the policy enforcement layer (e.g., the virtual networking infrastructure). In future work, we will investigate the automation of this step using, for example, the IETF policy model [91] and various graph-based mechanisms from the literature. We will also investigate different techniques for resolving conflicting policies [33, 17, 32, 38].

3.2 Virtual Networking Infrastructure

Virtual networking (VNET) technologies enable the seamless interconnection of VMs that reside on different physical hosts as if they were running on the same machine. In our TVD framework, we employ multiple technologies, including virtual switches, Ethernet encapsulation, VLAN tagging, and VPNs, to virtualize the underlying network and securely group VMs that belong to the same TVD. A single private virtual network is dedicated to each TVD, and network separation is ensured by connecting

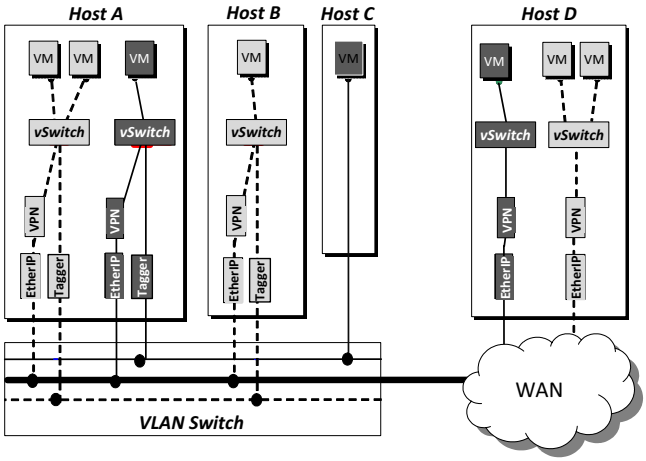

Figure 3.2: General vSwitch Architecture.

the VMs at the Ethernet level. Logically speaking, we provide a separate "virtual infrastructure" for each TVD in which we control and limit the sharing of network resources (such as routers, switches) between TVDs. This also provides the TVD owner with the freedom to deploy a wide range of networking solutions on top of the TVD network infrastructure. Network address allocations, transport protocols, and other services are then fully customizable by the TVD owner and work transparently as if the VMs were in an isolated physical network. To maintain secrecy and confidentiality of network data (where necessary), network communication is established over encrypted VPN tunnels. This enables the transparent use of untrusted networks between physical hosts that contain VMs within the same TVD to provide a seamless view of the TVD network.

In this section, we introduce the technologies we use to implement a security-enhanced VNET infrastructure for TVD owners. The concept of virtual switching is central to our architecture, which is then protected by existing VPN technologies that provide data confidentiality and integrity where needed. The VNET infrastructure acts as the local enforcer of VNET policies. As described in Section 2.3.2, these policies are based on the high-level TVD policies and translated into network configurations by the TVD Proxy. The Proxy then deploys the whole VNET infrastructure with respect to the translated configuration.

3.2.1 Virtual Switching

The *virtual switch* (vSwitch) is the central component of the virtual networking infrastructure and operates similarly to a physical switch. It is responsible for network virtualization and isolation, and enables a virtual network to span multiple physical hosts. To do so, the vSwitch uses EtherIP [52] and VLAN tagging [49] to insert VLAN membership information into every network packet. The vSwitch also implements the necessary address-mapping techniques to direct packets only to those machines that host member VMs. Virtual switches provide the primitives for implementing higher-level security policies for networking and are configured by the higher-level TVD management layer.

Figure 3.2 illustrates an example architecture in which physical machines host multiple VMs with different TVD memberships (the light and dark shades indicate different TVDs). Hosts A, B, and D are virtualized machines, whereas Host C is non-virtualized. Furthermore, Hosts A, B, and C reside on the same LAN, and thus can communicate directly using the trusted physical infrastructure without further protection (e.g., traffic encryption). For example, the *light* VMs hosted on Hosts A and B are inter-connected using the local VLAN-enabled physical switch. In this case, the physical switch separates the TVD traffic from other traffic passing through the switch using VLAN tags. Similarly, the *dark* VMs hosted on Host A and the non-virtualized Host C are seamlessly inter-connected using the local switch. In contrast, connections that require IP connectivity are routed over the WAN link. The WAN cloud in Figure 3.2 represents the physical network infrastructure able to deal with TVD-enabled virtual networks; it can include LANs with devices capable of VLAN tagging and gateways to connect the LANs to each other over (possibly insecure) WAN links. For connections that traverse untrusted medium, we employ EtherIP encapsulation to denote TVD membership and additional security measures (such as encryption) to ensure compliance with the confidentiality and integrity requirements.

3.2.2 Virtual Private Networking

In Figure 3.2, VMs hosted on Host D are connected to the other machines over a WAN link. A practical setting in which such a connection might exist would be an outsourced remote resource connected to the local data center through the Internet. As an example, lightly shaded VMs on Host D connect to the lone VM on Host B over this untrusted link. In this setting, we use a combination of EtherIP encapsulation and VPN technology to ensure the confidentiality and integrity of the communication. To do so, we use point-to-point VPN tunnels with OpenVPN that are configured via the TVD Proxy from the TVD policies. This enables reconfiguration of the topology and the involved VPNs within a TVD from a single administration point, the TVD Master.

TVD policies distributed from the TVD Master to the TVD Proxy also include the secret key for the VPN along with other VPN-specific settings. On a physical host, the VPN's endpoint is represented as a local virtual network interface (vif) that is plugged into the appropriate vSwitch controlled by the TVD Proxy. The vSwitch then decides whether to tunnel the communication between VMs, and if so, uses the VPN module to establish the tunnel and access the VPN secret for traffic encryption and decryption.

3.3 Virtual Storage Infrastructure

We focus on a simplified security management of virtualized storage. Broadly speaking, storage virtualization abstracts away the physical storage resource(s). It is desirable to allow a storage resource to be shared by multiple host computers, and also to provide a single storage device abstraction to a computer irrespective of the underlying physical storage, which may be a single hard disk, a set of hard disks, a Storage Area Network (SAN), etc. To satisfy both requirements, storage virtualization is typically done at two levels. The first level of virtualization involves aggregating all the (potentially heterogeneous) physical storage devices into one or more virtual storage pools. The aggregation allows more centralized and convenient data management. The second level of virtualization concerns the unified granularity (i.e., blocks or files) at

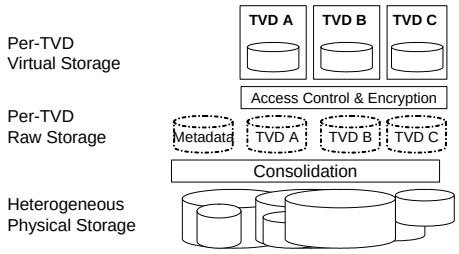

Figure 3.3: Security Enforcement for Virtualized Storage.

which data in each pool is presented to the higher-level entities (operating systems, applications, or VMs).

Figure 3.3 shows our storage security enforcement architecture, in which existing heterogeneous physical storage devices are consolidated into a joint pool. This virtual storage pool is then subdivided into raw storage for each TVD. Each raw storage volume has an owner TVD that determines its policy (indicated by the labels TVD A, TVD B, and TVD C at the per-TVD raw storage layer in the figure). In addition, when a volume shall be shared among multiple TVDs, there is also a set of member TVDs associated with it. The access control and encryption layer helps enforce the storage-sharing policy defined by the owner TVD, e.g., enforcing read, write, create, and update access permissions for the member TVDs. This layer is a logical layer that in reality consists of the virtual storage managers (part of the security services) located on each physical platform. The virtual storage manager on each physical platform is responsible for enforcing the owner TVD's storage security policies (see Section 2.3.3) on these volumes. If a certain intra-TVD security policy requires confidentiality and does not declare the medium as trusted, the disk is encrypted using a key belonging to the owner TVD.[1] If conditions for (re-)mounting a disk have been defined, the disk is also encrypted and the key is sealed against the TCB while including these conditions into the unsealing instructions. The policy and meta-data are held on a separate raw volume that is only accessible by the data center infrastructure.

An administrator of a domain may request that a disk be mounted to a particular VM in a particular mode (read/write). In Xen, the disk is usually mounted in the management machine Domain-0 as a *back-end device* and then accessed by a guest VM via a *front-end* device. The virtual storage manager on the platform validates the mount request against the policies of both the TVD the VM is part of and the owner TVD for the disk. Once mounted, appropriate read-write permissions are granted based on the flow control policy for the two TVDs, e.g., read access is granted only if the policies specified in the disk policy matrix allow the VM's TVD such an access to the disk belonging to the owner TVD.

3.4 TVD Admission Control

When a VM is about to join a TVD, different properties will be verified by the local TVD Proxy to ensure that policies of all the TVDs that the VM is currently a member of as well as of the TVD that it wants to join are not violated. If the verification is successful, then the VM will be connected to that TVD. The TVD admission control

[1] For efficiency reasons, we currently do not provide integrity protection.

protocol is the procedure by which the VM gets connected to the TVD. In the case of a VM joining multiple TVDs, the admission control protocol is executed for each of those TVDs. We now describe the steps of the protocol.

We assume that the computing platform that executes the VM provides mechanisms that allow remote parties to convince themselves about its trustworthiness. Example mechanisms include trusted (authenticated) boot and the remote attestation protocol (see Section 4.2) based on TPM technology.

TVD Proxy Initialization Phase: To allow a VM to join a TVD, the platform hosting the VM needs access to the TVD policy, and upon successful admission, to TVD secrets, such as the VPN key. For this purpose, TVD Proxy services are started on the platform for each TVD whose VMs may be hosted. The TVD Proxy can be started at boot time of the underlying hypervisor, by a system service (TVD Proxy Factory), or by the VM itself, as long as the TVD Proxy is strongly isolated from the VM.

Pre-Admission Phase: When a VM wants to join a TVD that is going to be hosted on the platform for the first time, the TVD Master has to establish a trust relationship with the platform running the VM, specifically with the TVD Proxy. We call this step the *pre-admission* phase. It involves the establishment of a trusted channel (see Section 4.3) between the TVD Master and the TVD Proxy (or the TVD Proxy Factory). The trusted channel allows the TVD Master to verify the integrity of the TVD Proxy (Factory) and the underlying platform. After the trusted channel has been established and the correct configuration of the Proxy has been verified, the TVD Master can send the TVD policies and credentials (such as a VPN key) to the TVD Proxy.

Admission Control Phase: The Compartment Manager (part of the platform security services shown in Figure 3.1) is responsible for starting new VMs. The Compartment Manager loads the VM configuration and enforces the security directives with the help of the Integrity Manager (also part of the platform security services shown in Figure 3.1). The security directives may include gathering the VM state information, such as the VM configuration, kernel, and disk(s) that are going to be attached to the VM.

If the VM configuration states that the VM should join one or more TVDs, then the Compartment Manager interacts with the corresponding TVD Proxy(ies) and invokes TPM functions to attest the state of the VM. The TVD Proxy verifies certain properties before allowing the VM to join the TVD. More concretely, the TVD Proxy has to ensure that

- the VM fulfills the integrity requirements of the TVD;
- the information flow policies of all TVDs the VM will be a member of will not be violated;
- the VM enforces specific information flow rules between TVDs if such rules are required by the TVD policy, and that
- the underlying platform (e.g., the hypervisor and attached devices) fulfills the security requirements of the TVD.

Platform verification involves matching the security requirements with the platform's capabilities and mechanisms instantiated on top of these capabilities. For example,

suppose that data confidentiality is a TVD requirement. Then, if hard disks or network connections are not trusted, additional mechanisms, such as block encryption or VPN (respectively), need to be instantiated to satisfy the requirement.

TVD Join Phase: If the VM and the provided infrastructure fulfill all TVD requirements, a new network stack is created and configured as described in Section 3.2. Once the Compartment Manager has started the VM, it sends an attach request to the corresponding TVD vSwitch. Once the VM is connected to the vSwitch, it is a member of the TVD.

Chapter 4

Background and Related Work

In order to put our work in context we survey key concepts that underlie our approach. Section 4.1 presents the TVD concept, which can be thought of as a virtualization of today's security zones while making security requirements explicit. Section 4.2 describes *Trusted Computing* concepts. The core of this concept is a security hardware device called *Trusted Platform Module* that guarantees certain security functionalities in spite of attacks. We finally survey related work on trusted channels in Section 4.3 and on secure virtual networking in Section 4.4.

4.1 Overview of Trusted Virtual Domains

Bussani *et al.* [16] introduced the concept of TVDs. A Trusted Virtual Domain consists of a set of distributed Virtual Processing Elements (VPEs), storage for the VPEs, and a communication medium interconnecting the VPEs [16, 62, 45]. The TVD provides a policy and containment boundary around those VPEs. VPEs within each TVD can usually exchange information freely and securely with each other. At the same time, they are sufficiently isolated from outside VPEs, including those belonging to other TVDs. Here, isolation loosely refers to the requirement that a dishonest VPE in one TVD cannot exchange information (e.g., by sending messages or by sharing storage) with a dishonest VPE in another TVD, unless the inter-TVD policies explicitly allow such an exchange. There is a TVD *infrastructure* (for each TVD) that provides a unified level of security to member VPEs, while restricting the interaction with VPEs outside the TVD to pre-specified, well-defined means only. Unified security within a virtual domain is obtained by defining and enforcing *membership requirements* that the VPEs have to satisfy before being admitted to the TVD and for retaining membership. Each TVD defines rules regarding information exchange with the outside world, e.g., restrictions regarding in-bound and out-bound network traffic.

Figure 1.1 shows customer VMs as VPEs belonging to TVD_1 spanning two platforms (contained in the dashed boxes). The Master (TVD1 Master) and Proxy components (Proxy1 on each platform) are part of the TVD infrastructure, which we describe in detail in Section 3.1. The TVD Master is the orchestrator of the TVD deployment and configuration. There is one TVD Proxy for each platform hosting VMs belonging to that TVD. If the platform hosts VMs belonging to multiple TVDs, then there are multiple TVD proxies on that platform, one per TVD. The TVD Proxy on a platform is configured by the TVD Master and can be thought of as the local TVD policy en-

forcer. VMs belonging to the same TVD can usually exchange information freely with each other unless restricted by VM-level policies. For example, traffic originating from VM_{A1} or VM_{A2} on Host A is routed to VM_{Bi} ($i = 1, \cdots, 4$) on Host B without any restrictions. Information exchange among TVDs can be allowed; however, it is subject to the network and storage policies stated by each TVD Master and locally enforced by each TVD Proxy.

4.2 Trusted Computing – The TCG Approach

It is important to have reliable mechanisms for a system to reason and verify the trustworthiness (i.e., compliance with a certain security policy) of a peer endpoint (local or remote). A recent industrial initiative towards realizing such a mechanism was put forward by the *Trusted Computing Group* (TCG) [113], a consortium of a large number of IT enterprises that proposes a new generation of computing platforms that employs both supplemental hardware and software (see, e.g., [81, 100]). The TCG[1] has published several specifications on various concepts of trusted infrastructures [121].

The Trusted Platform Module The core component the TCG specifies is the *Trusted Platform Module* (TPM). Currently, the widespread implementation of the TPM is a small tamper-evident chip[2] that implements multiple *roots-of-trust* [122, 120], e.g., the root-of-trust for reporting and the root-of-trust for storage. Each root-of-trust enables parties, both local and remote, to place trust on a TPM-equipped platform that the latter will behave as expected for the intended purpose. By definition, the parties trust each root-of-trust, and therefore it is essential that the roots-of-trust always behave as expected. Given that requirement, a hardware root-of-trust – especially one that is completely protected from software attacks and tamper-evident against physical attacks, as required by the TPM specification – is assumed to provide a better protection than software-only solutions.

Attestation and Integrity Verification The Trusted Computing features we leverage in this paper are protection of keys, secure recording of integrity measurements, attestation, and sealing. Integrity verification mechanisms enable a remote party to verify whether system components conform to certain security policies. *Measurement* of a component involves computing the SHA-1 hash of the binary code of that component. In particular, each software component in the Trusted Computing Base (TCB) is first measured and then its measurement recorded before control is passed to it. The hash values are then appended to a hash chain, which is kept in special protected registers called *Platform Configuration Registers* (PCRs), thus acting as accumulators for measurements. *Recording* a measurement means appending it to the hash chain by PCR extend operation[3]. The sequence of measured values are also stored in a *measurement log*[4], external to the TPM.

[1] TCG's claimed role is to develop, define, and promote open and vendor-neutral industry specifications for Trusted Computing, including hardware building blocks and software interface specifications across multiple platforms and operating environments.

[2] Many vendors already ship their platforms with TPMs (mainly laptop PCs and servers).

[3] Extending of PCR values is performed as follows: $PCR_{i+1} := \mathrm{SHA1}(PCR_i|I)$, with the old register value PCR_i, the new register value PCR_{i+1}, and the input I (e.g. a SHA-1 hash value).

[4] Since each PCR holds only the digest of (part of) the chain of trust, keeping the list of all measured values is required if afterwards, during the attestation process, a remote party wants to identify each measured component.

Attestation refers to the challenge-response-style cryptographic protocol for a remote party to query the recorded platform measurement values and for the platform to reliably report the requested values. The verifier first sends a challenge to the platform. The platform invokes the TPM_Quote command with the challenge as a parameter. The invocation also carries an indication of which PCRs are of interest. The TPM returns a signed *quote* containing the challenge and the values of the specified PCRs. The TPM signs using an Attestation Identity Key (AIK), whose public key is certified by a third party that the verifier trusts, called *Privacy CA* in TCG terminology. The platform then replies to the verifier with the signed quote along with the AIK public key certificate and the log information that is necessary to reconstruct the platform's configuration. Based on the reply, the verifier can decide whether the platform is in an acceptable state.

Sealing is a TPM operation that is used locally to ensure that a certain data item is accessible only under specific platform configurations reflected by PCR values. The *unsealing* operation will reveal the data item only if the PCR values at the time of the operation match the PCR specified values at the time of sealing.

A more general and flexible extension to the binary attestation is *property-based attestation* [96, 90, 69]: Attestation should only determine whether a platform configuration or an application has a desired property. However, our prototype is still using binary attestation.

In [47], the authors propose *semantic remote attestation* using language-based trusted VM to remotely attest high-level program properties. The general idea is to use a *trusted VM* (TrustedVM) that verifies the security policy of another virtual machine on a given host.

In [77], [78], and [79], the authors propose a software architecture based on Linux providing attestation and binding. The architecture binds short-lifetime data (e.g., application data) to long-lifetime data (e.g., the Linux kernel) and allows access to that data only if the system is compatible with a security policy certified by a security administrator.

4.3 Trusted Channels

The standard approach for establishing secure channels over the Internet is to use security protocols such as Transport Layer Security (TLS) [27] or Internet Protocol Security (IPSec) [63]), which aim at assuring confidentiality, integrity, and freshness of the transmitted data as well as authenticity of the endpoints involved. However, as mentioned before, secure channels do not provide any guarantees about the integrity of the communication endpoints, which can be compromised by viruses or Trojans. Based on security architectures that deploy Trusted Computing functionality, one can extend these protocols with integrity reporting mechanisms (e.g., the TLS extension proposed in [42, 9]). Such extensions can be based on binary attestation or on property-based attestation.

4.4 Secure Network Virtualization

Previous work on virtualizing physical networks can be roughly grouped into two categories: those based on Ethernet virtualization and those based on TCP/IP-level virtualization. Although both categories include a substantial amount of work, few of these

studies have an explicit focus on security.

A secure network virtualization framework was proposed by Cabuk *et al.* [18] for realizing the network flow aspects of TVDs. The focus of [18] is a security-enhanced network virtualization, which (1) allows groups of related VMs running on separate physical machines to be connected together as though they were on their own separate network fabric, and (2) enforces intra-TVD and inter-TVD security requirements such as confidentiality, integrity, and inter-TVD flow control. This has been achieved by an automatic provisioning of networking components such as VPNs, Ethernet encapsulation, VLAN tagging, and virtual firewalls.

A second concept for managing VLAN access has been proposed by Berger *et al.* in [14]. Both papers contain similar concepts for managing VLANs inside the data center with some differences. The work of Berger *et al.* has more of a focus on integrity assurance using Trusted Computing. The work of Cabuk et al [18] allows provisioning of secure virtual networking even if no VLAN infrastructure is present.

Part II

Building Blocks of the OpenTC Security Architecture

Chapter 5

Policy Enforcement and Compliance Proofs for Xen Virtual Machines

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5.1 Introduction

Hardware virtualization is enjoying a resurgence of interest fueled in part by its cost-saving potential in data centers. By allowing multiple virtual machines to be hosted on a single physical server, virtualization helps improve server utilization, reduce management and power costs, and control the problem of server sprawl.

We are interested in the security management of virtual machines, i.e., the protection, enforcement, and verification of the security of virtual machines. Security management is a non-trivial problem even in traditional non-virtualized environments. Security management of virtual machines (VMs) is even more complicated because the virtual machines hosted on a given physical server may belong to different virtual organizations, and as a result, may have differing security requirements. Protecting a VM against security attacks may be complicated by inadequate isolation of the VM from other VMs hosted on the same server. Verifying the security of a VM may be complicated by confidentiality requirements, which may dictate that the information needed for verification of a VM's configuration should not divulge configuration information of other co-hosted VMs.

We address two main problems relating to security management, particularly integrity management, of VMs: (1) protecting the security policies of a VM against modification throughout the VM's life cycle, and (2) verifying that a VM is compliant with specified security requirements. We describe a formal model that generalizes integrity management mechanisms based on the Trusted Platform Module (TPM) [120] to cover VMs (and their associated virtual devices) and a wider range of security policies (such as isolation policies for secure device virtualization and migration constraints for VMs). On TPM-equipped platforms, system compliance can be evaluated by checking TPM register values. Our model allows finer-grained compliance checks by handling policies that can be expressed as predicates on system log

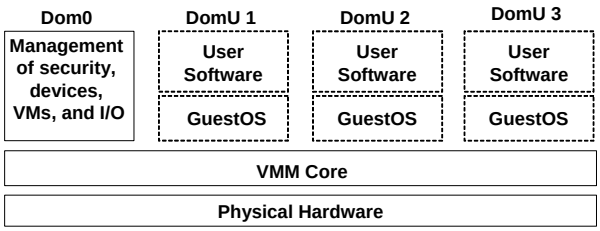

Figure 5.1: Xen virtual machine architecture

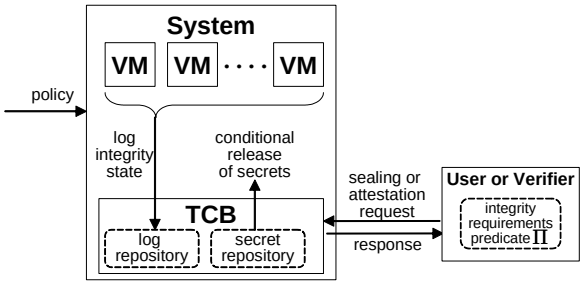

Figure 5.2: System model for integrity management

entries. Verifying compliance involves showing that the system integrity state, as reflected by secure write-only logs, satisfies certain conditions. We build on previous work by others [39, 45, 101, 104, 13] who have used the Trusted Platform Module (TPM) [120] to protect the integrity of the core virtual machine monitor (VMM) and to reliably isolate VMs. Based on the formal model, we describe an integrity architecture called PEV (which stands for protection, enforcement, and verification) and associated protocols. The architecture incorporates integrity protection and verification as part of the virtualization software itself, and at the same time enhances its policy enforcement capabilities. We describe a prototype realization of our architecture using the Xen hypervisor [11]. We demonstrate the policy enforcement and compliance checking capabilities of our prototype through multiple use cases.

Our generalized integrity management mechanisms are both extensible and flexible. *Extensibility* means that it is possible to guarantee compliance even if new virtual devices are attached to the VMs. *Flexibility* means that the verifier is able to specify which aspects of the enforced security policies are of interest, and obtain only the information corresponding to those aspects for validation of system compliance.

5.2 Formal Integrity Model for Virtual Machines

Figure 5.2 shows our system model for integrity management. At a high level, the system consists of VMs and a TCB, and is configured through policies. The TCB periodically logs the integrity state of the rest of the system. The log repository contains a record of the integrity history of the system, and is *secure write-only*, i.e., log entries, once written, cannot be modified or removed by any entity in the rest of the system.

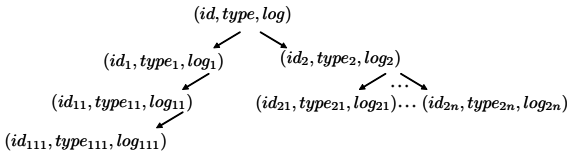

Figure 5.3: Tree $\mathcal{T}$ of log entries

The log data includes the list of software components, configuration parameters, policies, and any updates to them. The log contents are useful in evaluating compliance with those security properties that can be expressed as predicates on the contents. The compliance proof involves showing that correct policies and healthy policy enforcement mechanisms are in place. The TCB also provides conditional release of secrets, where the condition is expressed as predicates on the log data. That allows a sensitive data item and a condition to be stored such that the data item is released only if the log data satisfies the condition specified.

For flexibility and extensibility, the log data is stored in a tree structure instead of a monolithic log file. The log tree $\mathcal{T}$ is shown in Figure 5.3. Each tree node is a triple containing log data for one system component. To keep the tree size manageable, only those components that have an impact on the system's integrity or those that are of interest from an integrity verification point of view are represented in the log tree. A triple for a component k contains an identifier id_k, a component type $type_k$, and a vector log_k of log values. Sub-components are modeled as children of a node. The tree can be extended by adding or removing children nodes. For example, the addition of a new virtual device to a VM can be easily reflected in the log tree by adding a new node as a child of the sub-tree that corresponds to the VM.

The integrity requirements of a user or verifier are modeled by $\Pi(\mathsf{p}(\mathcal{T}))$, where Π is a predicate and $\mathsf{p}()$ is a projection function. We introduce the notion of a *projection function*, denoted by $\mathsf{p}()$, to model the specific aspects of the system's integrity state that is of interest to a user or verifier. For example, a verifier may be interested only in a disk's access control list and not the actual disk contents. When applied on the log tree, the function returns a subset of the tree nodes and a subset of the vector elements from the log vector of each node. Formally, $\mathsf{p}(\mathcal{T}) = \{l_k\}$, where $l_k \in log_k$, and $(id_k, type_k, log_k) \in \mathcal{T}$.

We now use our formal model to generalize TPM-based integrity protection and verification. We also enhance our model by adding access control to the log contents.

5.2.1 Generalized Sealing to Protect Integrity

A TPM-equipped system can *seal* a data item, i.e., the system can encrypt the data item and bind it to the system configuration prevalent at the time of sealing. The system configuration is reflected by the contents of a specified subset of the TPM's PCRs. The data item may be a key generated by the TPM itself or something generated outside the TPM. Decryption of the data item, called *unsealing*, is possible only when the system configuration (reflected by the contents of the same subset of PCRs) is the same as that at the time of sealing.

We generalize sealing for protecting the integrity of a sensitive data item d by making d inaccessible to the system (or some component) unless specified integrity re-

quirements are met. We use two operations, seal and unseal, to model the concept of generalized sealing. Let $\mathcal{T}_x$ denote the log tree at time t_x. The seal operation, performed at time t_s, takes as input the data item d, a projection function $\mathsf{p}()$, a sealing predicate Π, and the public part K_p of an encryption key K. The operation logs $\mathsf{p}()$ and Π, and encrypts d using K_p to produce the encrypted output e. Thus, the contents of $\mathcal{T}_s$ include $\mathsf{p}()$ and Π. The unseal operation, performed at time t_u (where $t_u > t_s$), takes as input e and $\mathcal{T}_u$, and outputs d iff the condition $\Pi(\mathsf{p}(\mathcal{T}_u))$ holds. In other words, the private part of the key K used for decrypting e is revealed iff the condition holds. Here, $\mathsf{p}()$ and Π are retrieved from the log. A simple sealing predicate may just compare the result of $\mathsf{p}(\mathcal{T}_u)$ with a reference value (e.g., $\mathsf{p}(\mathcal{T}_s)$). A more complex predicate may extract the high-level properties of the system from $\mathsf{p}(\mathcal{T}_u)$ and compare them with desired properties (similar to property-based attestation [90, 20, 47]).

One can easily see that our generalized sealing concept covers the special case of TPM sealing. For TPM sealing, $\mathcal{T}_u$ consists of the values in the PCRs; the projection function $\mathsf{p}()$ specifies the subset of PCRs whose values are of interest for assessing the system's integrity; the sealing predicate Π checks whether their values at the time of unsealing are the same as at the time of sealing.

5.2.2 Generalized Attestation to Verify Integrity

A TPM-equipped system can use the TPM to engage in a challenge-response style cryptographic protocol, called *attestation*, with a verifier. The protocol allows the verifier to query and reliably obtain the measurement values for the system stored in the PCRs of the TPM. Reliable reporting of the measurement values is due to the signing of the values by the TPM, which is trusted by the verifier. Based on these values, the verifier can assesses the integrity state and the trustworthiness of the system.

We generalize attestation so that the verifier can specify which aspects of the system's integrity state are of interest to her. In our model, the attestation operation $\mathsf{attest}()$ obtains as input a challenge c, an attestation predicate Π, a projection function $\mathsf{p}()$, and a secret key K_s. The operation outputs a signed message $\mathsf{sign}_{K_s}(\mathsf{f}(\mathsf{p}(\mathcal{T})), c)$.

Our attestation operation is a generalization of both binary and property-based attestation [90, 20, 47]. For binary attestation, the predicate Π is simply the identity function, i.e., $\Pi(x) = x$, and the result of attestation is simply the signature on the result of the projection function applied on the log tree. TPM attestation is a special case of binary attestation in which $\mathcal{T}$ simply consists of the values in the PCRs and the projection function $\mathsf{p}()$ specifies a subset of PCRs. For property-based attestation, the predicate Π extracts high-level properties from the result of the projection function applied on the log tree.

Whereas previous works such as the Integrity Measurement Architecture (IMA) of of Sailer *et al*. [104] provide a good way of checking the hash of software binaries, our generalized attestation enables better assessment of the run-time behavior of the system. In this respect, our model has goals similar to those of Haldar *et al*. [47]. However, unlike Haldar *et al*. who focus on attesting the behavior of a software application, our model has a focus on VMs and virtual devices. For example, our attestation operation enables a verifier to check the number and type of VMs running on the system. Because of their reliance on the Java virtual machine which runs on top of an operating system, their TCB includes the operating system. In contrast, our TCB includes only the VMM and underlying system layers, and is much smaller than theirs.

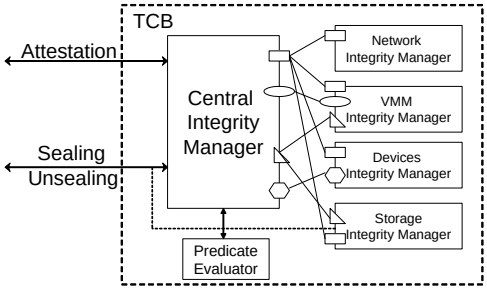

Figure 5.4: Architecture for integrity protection and verification

5.2.3 Access Restriction

The integrity of certain aspects of the system (such as the VMM) may be important to multiple users. Conversely, certain aspects of the system may be confidential to one or more users, e.g., the state of a particular VM may be verified only by the users of that VM. Hence, it is important that attestation and sealing be applied not directly on the system state, but on appropriate projections of the state. Furthermore, if a state that is relevant for integrity verification contains information about multiple users, it should be possible to prove integrity without revealing the actual state. We formalize such requirements using two concepts: *access restriction specification* and *projection assessment function*.

Given a set of users U and a log tree $\mathcal{T}$, an access restriction is specified by a function $\mathsf{r}()$ that assigns a subset of U to each vector element in each node of the tree. The subset assigned to a given vector element in a given node is called the *access control list* (ACL) for that element. Despite the potentially large number of nodes in the log tree, ACLs can be efficiently implemented by attaching ACLs only to some nodes and vector elements. ACLs of children nodes may be derived through inheritance of the parent node's ACL. Scoping rules may be used to apply an ACL to multiple vector elements of a given node.

A projection assessment function can determine whether a given projection conforms to or violates access restrictions. A projection $\mathsf{p}(\mathcal{T})$ applied by a user $u \in U$ conforms to the access restriction specification $\mathsf{r}()$ if the output only contains vector elements in which u was contained in the ACL. Any predicate Π for attestation or sealing can be applied on such a projection without violating the access restrictions. If the projection does not conform to $\mathsf{r}()$, then prior to applying the predicate, an *access filter* is used to hide those parts of $\mathsf{p}(\mathcal{T})$ that u is not authorized to see.

5.3 The PEV Integrity Architecture

Figure 5.4 shows the PEV architecture for protecting, enforcing, and verifying the integrity of VMs and virtual devices. There is a *central integrity manager* and *component integrity managers* that are associated with individual system components such as storage, VMM, networking, and other devices. Each component integrity manager is responsible for the part of the log tree corresponding to the component. For example, the storage integrity manager is responsible for maintaining the storage sub-tree of the

system log tree $\mathcal{T}$. Hereafter, we refer to the central integrity manager as *the* Integrity Manager.

The Integrity Manager has a *master plug-in module* for each log projection function that needs to be implemented. The module obtains state information about various aspects of the system that may be of interest to a potential verifier or user by invoking the appropriate *component plug-in modules* and aggregating their outputs. A component plug-in module is part of the component integrity manager and reveals particular aspects of the component's integrity that are relevant for the projection function.

In Figure 5.4, the various master plug-in modules are attached to the Integrity Manager are shown using different geometrical shapes (ovals, hexagons, triangles, and rectangles). For example, the triangular plug-in module measures certain aspects of system storage and the VMM, as indicated by the presence of triangular component plug-in modules in the Storage Integrity Manager and VMM Integrity Manager. On the other hand, the hexagonal plug-in module measures only certain aspects of system devices. Each plug-in module has a unique identifier. The mapping between each plug-in identifier and the functionality provided by the corresponding plug-in module is made publicly available (e.g., through a naming service or a published table).

5.3.1 Sealing/Unsealing Protocol

At the time of sealing, the user provides the following inputs:

Data The data item to be encrypted during sealing and to be revealed later only if certain conditions are met.

Key The sealing key whose public part is used for encrypting the data at the time of sealing, and whose private part is revealed only if the unseal operation completes successfully.

Identifier(s) of Plug-in Module(s) By listing the identifiers of plug-in modules, a user can choose what aspects of the system's integrity state are to be checked prior to revealing the private part of the sealing key.

Predicate The predicate specifies user-defined conditions that the system's integrity state must satisfy at the time of unsealing in order for the private part of the sealing key to be revealed.

Our sealing protocol requires the log projection functions (described in Section 5.2.1) to be implemented as plug-in modules as part of the TCB. The key used for encrypting the sensitive data item is sealed away against the state of the TCB and a hash of the user-specified projection functions and sealing predicates. The Integrity Manager stores the state of the TCB in PCRs that cannot be reset and the hash in a resettable PCR (say PCR_i). This ensures that the TCB is aware of the conditions to be satisfied before the key can be revealed to the user. To perform the unseal operation, the TCB has to ensure that PCR_i still contains the hash of the user-specified projection function and sealing predicates. Then, the unseal operation reveals the key to the Integrity Manager. The Integrity Manager then invokes the *predicate evaluator* module (Figure 5.4) to check whether the sealing predicates (evaluated on the output of the log projection function) are indeed satisfied. If that is the case, then the Integrity Manager reveals the key to the user.

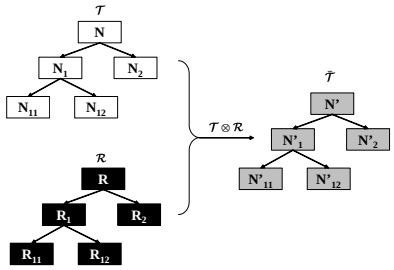

Figure 5.5: Enforcing access restrictions on system state

The flexibility of our sealing protocol is due to the fact that arbitrarily complex conditions to reveal the sealed key can be coded as plug-in modules. The extensibility arises from the fact that new plug-in modules covering the integrity state of newly added VMs or virtual devices can be easily added to the TCB.

5.3.2 Attestation Protocol

The verifier initiating the attestation protocol provides as input a challenge (to ensure freshness) and the identifier(s) of plug-in module(s) that are relevant to evaluating system compliance with the verifier's integrity requirements. The flexibility of our attestation protocol relies on the verifier being able to attest the TCB and requires the log projection functions (described in Section 5.2.1) to be implemented as plug-in modules as part of the TCB. The extensibility of our attestation protocols relies on the ability to add new plug-in modules for new aspects of the system's integrity state that the verifier may be interested in.

5.3.3 A Blinding Technique For Enforcing Access Restrictions

Figure 5.5 shows a simple *blinding* technique that uses a *commitment scheme* to enforce access restrictions on the log tree. Cryptographic commitment schemes [44] generally consist of two phases. The first phase, called *commit phase*, is used to make a party *commit* to a particular value while hiding that value from another party. It is only in the second phase, called *reveal phase*, that the value is *revealed* to the second party. Any commitment scheme guarantees that (a) the committed value cannot be obtained by the second party before the reveal phase, and (b) the second party can detect whether the value revealed is indeed the same value that was committed to in the first phase.

For simplicity, we consider blinding at the granularity of log tree nodes instead of at the granularity of log vector elements in the tree nodes, i.e., the access restriction specification $\mathsf{r}()$ assigns a subset of U to each node of the tree. A *random* tree $\mathcal{R}$ is bound to the original log tree $\mathcal{T}$ through a *multi-bit commitment scheme* to give the blinded log tree $\bar{\mathcal{T}}$. $\mathcal{R}$ is a tree that has the same structure as that of $\mathcal{T}$ and whose nodes are random numbers. Existing commitment schemes such as the one by Damgard et al. [26] or those based on one-way hash functions can be used for this purpose.

In a TPM-equipped system, logging is done by extending the PCRs with the measurement values. For blinding, it is the nodes of $\bar{\mathcal{T}}$ that are actually logged. This

means that instead of doing the normal TPM_extend(n), a TPM_extend(r $\otimes$ n) is done, where n is a node of $\mathcal{T}$, r is a node of $\mathcal{R}$, and $\otimes$ denotes the commitment operation used for hiding n until the reveal phase.

A projection function p() that conforms to the access restrictions can be realized as follows: when invoked at the request of user u, p() reveals $\bar{\mathcal{T}}$ and only those nodes in $\mathcal{T}$ that contain u in their respective ACLs. Thus, p() implements the reveal phase of the multi-bit commitment scheme and reveals only those nodes in $\mathcal{T}$ that u is authorized to access. Due to the guarantees of the commitment scheme, the system cannot invent arbitrary values for the nodes in $\mathcal{T}$ without being detected by the user.

As a result of the blinding technique described above, any user u knows that all components that have any effect on system integrity have been taken into consideration in the system log tree; in addition, for those components that it is authorized to access, u can check whether they indeed have the acceptable configuration and state value, by comparing with its own reference values that may be provided and certified by a trusted third party. In particular, if the ACL for the root node is U (i.e., all users can access the root node), then any user can verify overall system integrity (just from the value of *root*($\mathcal{T}$)) without knowing the exact configuration of any individual component in the system.

Our approach of using commitment schemes for blinding suffers from the disadvantage that two colluding verifiers can learn the values revealed to the other. Alternate schemes based on zero-knowledge proofs or deniable signatures need to be investigated to overcome this disadvantage.

5.4 Realization using Xen and Linux

Figure 5.6 shows an example implementation of our PEV architecture with the Xen hypervisor using Linux for Dom0. The main components of our implementation are the Compartment Manager (CM), Integrity Manager (IM), and the Secure Virtual Device Manager (SVDM). All components are implemented in Dom0. The CM is responsible for the VM life cycle management. As the sole entry point for user commands, it communicates directly with the hypervisor and orchestrates the IM and the SVDM. Table 5.1 shows the mapping between concepts in our formal model and their realization in our Xen prototype. XSLT is a language for transforming one XML document into another XML document [6]. We assume that the XSLT interpreter is part of the TCB.

The getCurrentState() function of the CM returns the current state of the physical machine, which includes the list of hosted VMs, their status (active, suspended, or hibernating), VM ownership information (e.g., the virtual organization to which a VM belongs), the amount of free memory available, etc. Using the result of the function, a verifier can decide whether the physical machine satisfies the integrity requirements for performing certain actions (e.g., starting a new VM belonging to a particular virtual organization).

The IM in our Xen prototype has a storage integrity plug-in (SIP) for measuring various disk images and files. The IM also has an Attestation & Sealing module (ASM) that interfaces with the TPM for executing the sealing and attestation protocols (described in Sections 5.3.1 and 5.3.2) as well as for invoking normal TPM operations, such as TPM_Quote. The ASM invokes normal TPM operations through the TPM Software Stack (TSS) [116], which is the standard API for accessing the functions of the TPM.

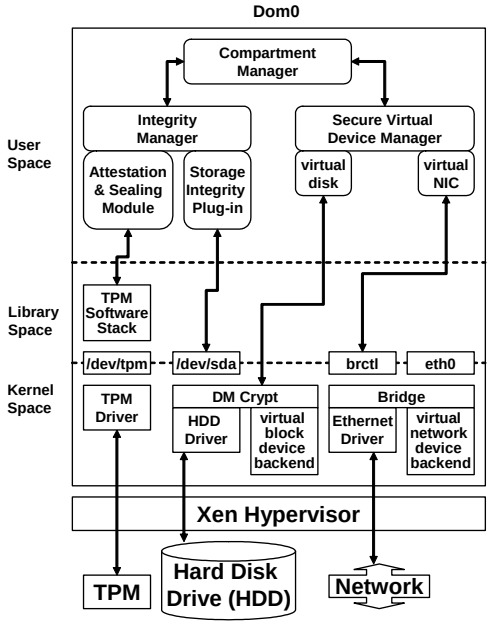

Figure 5.6: Realization using Xen and Linux

Model		Xen-based Prototype
Projection	p()	component measurement plug-in
Predicate	Π	XSLT stylesheet
Access Filter		XSLT stylesheet

Table 5.1: From model to implementation

The SVDM is responsible for managing virtual devices such as virtual hard disks, virtual block devices, virtual network devices, and virtual TPMs. The service offered by the SVDM is realized through multiple specialized low-level component plug-in modules, one for each virtual device. Figure 5.6 shows two plug-ins in our Xen prototype. One is for managing the virtual (encrypted) hard disk and the other one is for managing the virtual network interface card (NIC).

In Dom0, secure device virtualization is implemented in the kernel space. Tasks such as configuring virtual devices are done through the SVDM in the user (or application) space. The SVDM manages the security properties of devices. For example, a secure hard disk is implemented by means of the *DM Crypt* loopback device. Similarly, network virtualization is done by providing virtual NICs for the VMs and *bridging* these virtual NICs to the physical NIC. Security for networks has two aspects. Topology constraints define which VM is allowed to connect to which sub-network(s). In addition, confidentiality requirements dictate which connections need to be encrypted.

Secure management of virtual devices is a complex task. For example, there are multiple steps involved in starting a virtual hard-disk drive. First, a policy-based check of the state of the physical machine is done based on the results of getCurrentState() function. Depending on the logic implemented by the corresponding plug-in, that check may include verifying the measurements of the hypervisor, binary disk, and the Dom0 image. Then, the virtual hard-disk is attached with credentials and connected to a loop device (/dev/loop). The virtual hard-disk may be encrypted, for example, with a sealing key that is made available only if the platform is in a certain state. The decryption of the virtual hard-disk image is done using the Linux hard-disk encryptor. After decryption, the device file that gives access to the decrypted image is connected to the front-end. Similar policy-based checks may be done when starting other virtual devices. For example, before starting a virtual network device, policies may stipulate that the VM must be in some acceptable state and outside firewalls must be configured correctly.

5.5 Use Cases

In this section, we describe a few examples of how the components introduced in Section 5.4 interact for integrity protection, enforcement, and verification purposes. We assume that the core TCB (including Xen and Dom0 Linux) has been measured at start-up time. Additional services may need to be measured based on policy. The measurement can either be done by a trusted boot loader such as TrustedGRUB [4] (which measures the entire boot image) or by a more fine-grained approach such as Sailer *et al.*'s IMA [104].

5.5.1 TPM-based Attestation on a VM Disk

Figure 5.7 shows the component interactions for attesting the current state of the TCB and the status of a VM's disk image. The user/verifier interacts with the CM through the attestationRequest call with an *attestation descriptor* and *user credential* as parameters (step 1). The attestation descriptor is an XML structure that describes what aspect of the system's integrity state the verifier wants attested. In other words, the attestation descriptor is how the verifier chooses the log projection function suitable

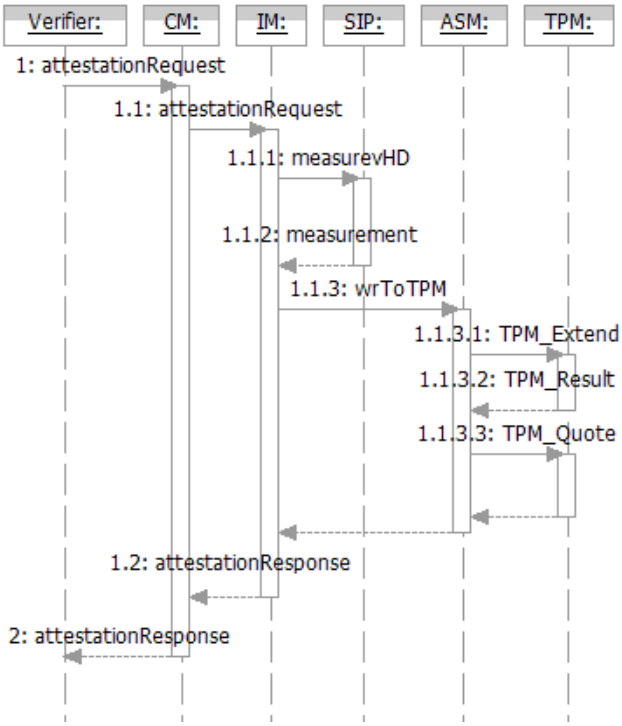

Figure 5.7: TPM-based attestation on a VM disk

```
<attestation-desc>
  <attestation type="tpm-based"
       challenge="0xaded..."
       aik="0xaada3..">
    <measurement-desc type="tpm">
        <measureTarget name="disk:/dev/sdb1"
           dest="PCR16"/>
    </measurement-desc>
    <attestTarget name="ALLPCRS"/>
  </attestation>
</attestation-desc>
```

Figure 5.8: Attestation descriptor in XML

for its purpose. As described before, projection functions are realized by a set of component plug-in modules. Some of these plug-in modules are *measurement plug-ins*, which not only return the relevant integrity states of the components but are also the ones measuring their integrity states in the first place. The attestation descriptor contains one or more *measurement descriptors*. Based on the measurement descriptors, the IM knows the exact set of measurement plug-ins to invoke.

Figure 5.8 shows an example attestation descriptor as a XML structure. It contains an `<attestation>` section, which defines the type of attestation (`tpm-based`) and the parameters needed for attestation (the TPM Attestation Identity Key or `AIK` and a `challenge`). Nested in the attestation descriptor is a measurement descriptor, which specifies a measurement target (`measureTarget`) and a destination (`dest`). The target indicates what is to be measured (in this case, a VM disk image), whereas the destination indicates where the result should be stored (in this case, the TPM's PCR number 16). The `<attestTarget>` defines the scope of the requested attestation (in this case, all PCRs).

Based on the user credential supplied, the CM checks whether the verifier has the right to request attestation of the system sub-states indicated by the attestation descriptor. The check is essentially a way of determining whether the requested projection is a projection that conforms to the access restriction specification; hence, it is useful in enforcing access restriction. If the check reveals that the verifier wants to have more attested than what he/she is allowed to, then the entire attestation request is denied. Otherwise, the CM forwards the request to the IM (step 1.1).

The IM extracts the measurement descriptor(s) from the attestation descriptor and delegates the measurement(s) to the appropriate plug-in(s). In our example, the IM invokes the measurevHD function at the SIP passing the measurement descriptor as a parameter (step 1.1.1). The plug-in completes the requested measurement and returns the measurement result back to the IM (step 1.1.2). Although step 1.1.2 might look like an unnecessary extra step, the indirection via the IM allows the measurement plug-ins to be written independent of the TPM or similar future devices that are indicated as `dest`.

The IM invokes the wrToTPM function at the ASM with the challenge, the `AIK`, the measurement result, and the destination PCR (step 1.1.3). The actual writing of the result into the PCR happens by the TPM_extend operation (step 1.1.3.1). Thereafter, a TPM_Quote gets created and returned to the ASM (steps 1.1.3.2 and 1.1.3.3). The ASM wraps the TPM_Quote into an attestationResponse and returns it to the IM. The

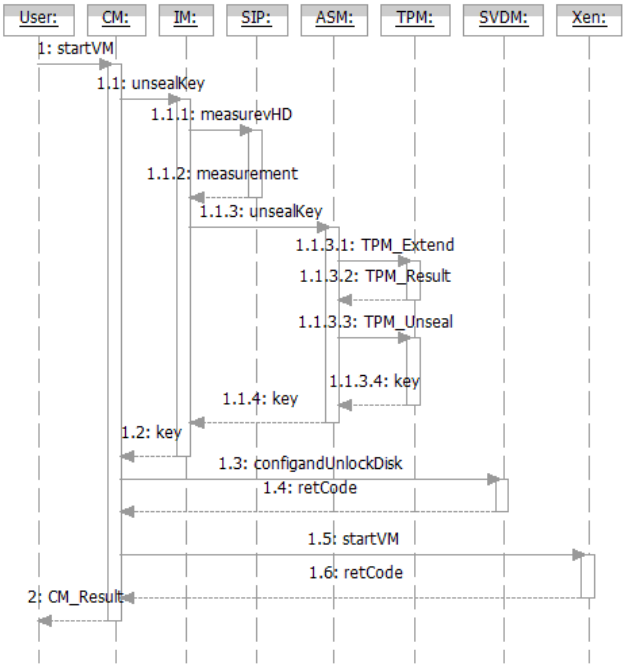

Figure 5.9: Creation of a VM with TPM-based sealing

attestationResponse includes not only the TPM_Quote but also the relevant log files. The IM returns the attestationResponse to the CM (step 1.2), which forwards it to the verifier (step 2).

A verifier can check the attestation result by recomputing a hash over the attestation targets (i.e., the relevant log files) specified in the attestationResponse and comparing the resulting hash with the hash in the PCR from the TPM_Quote.

The PCR in which the measurement result is stored will be reset after the attestation process has finished. Therefore, our prototype requires a TCG 1.2 compliant TPM, and the `dest` PCR has to be 16 or higher.

5.5.2 (Re-)Starting a VM with TPM-based Sealing

Figure 5.9 shows the component interactions for (re-)starting a VM with a sealed disk image. In this use case, we show how to enforce a policy that specifies that the key for decrypting the disk image be revealed only after measuring the disk image and only if the measurement value written into a specified PCR matches the value against which the key was sealed.

The user interacts with the CM through the startVM call to (re-)start the VM (step 1). After determining that the disk image has to be first decrypted through unsealing, the CM obtains the *sealing descriptor* that was given to it at the time of sealing. Like the attestation descriptor, the sealing descriptor also contains one or more measurement descriptors, which are used to let the IM know the exact set of measurement plug-in modules to invoke.

```
<sealing-desc>
  <sealing type="tpm-based" keyid="0x01">
    <measurement-desc type="tpm">
        <measureTarget
''          name="file:/xenimages/vm1image"
            desc="PCR16"/>
    </measurement-desc>
  </sealing>
</sealing-desc>
```

Figure 5.10: Sealing descriptor in XML

Figure 5.10 shows an example sealing descriptor as an XML structure. It contains an `<sealing>` section, which defines the type of sealing (`tpm-based`) and the parameters needed for unsealing (the identifier of the key protected by the TPM). Nested in the sealing descriptor there is a measurement descriptor, which specifies a measurement target (`measureTarget`) and a destination (`dest`). The target indicates what is to be measured (in this case, a VM disk image), whereas the destination indicates where the result should be stored (in this case, the TPM's PCR number 16).

The CM calls the IM interface unsealKey (step 1.1), passing the sealing descriptor as a parameter. The IM extracts the measurement descriptor from the sealing descriptor and calls the measurevHD interface of the SIP with the measurement descriptor (step 1.1.1). The plug-in reads the list of `measureTargets`, and accordingly measures the disk image. It returns a measurement result list to the IM (step 1.1.2). The IM calls the ASM, which handles TPM-related functions (step 1.1.3). The ASM writes the measurements to the TPM by invoking the TPM_Extend operation (step 1.1.3.1). Furthermore, the ASM performs the unsealing of the key requested by invoking the TPM_Unseal operation (step 1.1.3.3). If the `dest` PCR value matches the value at the time of sealing, then the disk is in the desired state and the unseal operation is successful (step 1.1.3.4); in that case, the ASM returns a key back to the IM (step 1.1.4), which in turn returns the key to the CM (step 1.2). In case the unseal operation fails, the ASM would return a failure. The CM calls the SVDM function configAndUnlock() to attach and unlock the disk (steps 1.3 and 1.4). Upon successful completion of that function, the CM instructs the Xen hypervisor to actually start the VM (steps 1.5 and 1.6).

For the sake of simplicity, Figure 5.9 does not show details of key handling such as loading a sealing wrapper key into the TPM.

5.5.3 Enforcement and Compliance Proofs for Information Flow Control

Consider, for example, the virtual network topology shown in Figure 5.11 with four virtual network zones. The topology shows the network of a company (which we shall call the *customer* company) connected to the Internet via a demilitarized zone (DMZ). The customer network is also connected to a *management network* that allows an outsourcing provider to manage the customer systems. The management network is not connected to the Internet.

An information flow-control matrix is a simple way of formalizing the system-wide flow-control objectives [18]. Figure 5.12 shows a sample matrix for the four virtual

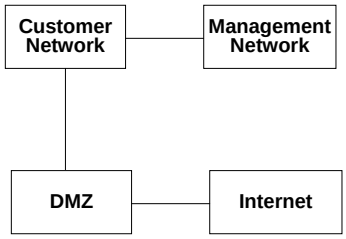

Figure 5.11: Virtual network topology

from/to	**Cust.**	**DMZ**	**Mgmt.**	**Internet**
Cust.	1	1	1	0
DMZ	1	1	0	1
Mgmt.	1	0	1	0
Internet	0	1	0	0

Figure 5.12: Flow control matrix

```
<flow-policy>
  <zone id="customer-net">
    <permit id="mgmt-net" />
    <permit id="dmz" />
  </zone> ...
</flow-policy>
```

Figure 5.13: Flow control policy in XML

network zones. Each matrix element represents a policy specifying the information flows permitted between a pair of network zones. The 1 elements along the matrix diagonal convey the fact that there is free information flow within each network zone. The 0 elements in the matrix are used to specify that there should be no information flow between two zones, e.g., between the management zone and the Internet.

In [18], we described a Xen-based prototype of a secure network virtualization architecture that is based on the concept of Trusted Virtual Domains. The architecture allows arbitrary network topologies connecting VMs. For example, different VMs on the same physical infrastructure may belong to different virtual network zones. Despite this, the architecture ensures the enforcement of policy-based information flow control. We can use the architecture for enforcing the policies shown in Figure 5.12.

By combining the Xen prototypes of our PEV architecture and our secure network virtualization architecture, it is possible to validate the configuration of the virtual networking subsystem on each host. The subsystem exports an XML version of its flow-control matrix, as shown in Figure 5.13. The network measurement plug-in outputs the XML structure of the flow-control policy, when invoked by the IM. By requesting attestation of the TCB and this policy, a verifier can obtain a compliance proof for the correct configuration of the virtual networking subsystem on a given host. At the

```
<xsl:template
  match="/flow-policy/zone[@id='customer-net']">
  <xsl:choose> <xsl:if
      test="count(*[@id='dmz'])=1
            and count(*[@id='mgmt-net'])=1">
      <true />
    </xsl:if>
  </xsl:choose>
</xsl:template>
```

Figure 5.14: XSLT condition

verifier, a XSLT stylesheet is used to perform further transformations on the XML file returned by the platform. The XSLT stylesheet is a concrete implementation of the attestation predicate Π (described in Section 5.2.2), which assesses whether the platform is trustworthy from the verifier's point of view. The result of the predicate will serve to convince the verifier that the policy in Figure 5.12 is the actual flow-control policy as enforced by the network subsystem. If access restriction is an important concern, the XML output from the plug-in modules may be first processed by an XSLT stylesheet that implements a access filter before passing it on to the verifier. In such a case, the stylesheet would be embedded in the platform TCB.

A user can also protect sensitive information (say, an encryption key) against access by an untrusted network configuration using a two-stage procedure. The first stage is sealing, in which the user has to specify the binary configuration of the TCB and conditions for checking whether a given network configuration is a trusted one. Figure 5.14 shows an XSLT script that encodes the condition that the customer network should be directly connected only to the DMZ and the management network of the outsourcing provider, but not to any other network. The input to the XSLT script is the XML policy that is output by the network measurement plug-in. The XSLT script is a concrete realization of the user-specified predicate Π in our formal model (Section 5.2). The user seals the key to both the state of the TCB and the value of a resettable PCR; the latter reflects the integrity of the XSLT script and the integrity of the plug-in identifier. The second stage is unsealing, in which the IM (i) obtains the result of the plug-in, (ii) applies the result as input to the XSLT script, (iii) extends the resettable PCR with the hash of the XSLT script and the network measurement plug-in identifier, and (iv) tries to unseal the actual key. For steps (iii) and (iv), the IM invokes the ASM. The TPM should only reveal the key if the TCB is correct and the XSLT evaluated to `<true/>` when executed on their output.

5.6 Conclusion

We introduced a formal model for managing the integrity of arbitrary aspects of a virtualized system and evaluating system compliance with respect to given security policies. Based on the model, we described an architecture, called PEV, for protecting security policies against modification, and allowing stakeholders to verify the policies actually implemented. We generalized the integrity management functions of the

Trusted Platform Module, so that they are applicable not just for software binaries, but also for checking security compliance and enforcing security policies. We described a prototype implementation of the architecture based on the Xen hypervisor. We also presented multiple use cases to demonstrate the policy enforcement and compliance checking capabilities of our implementation.

Chapter 6

Hierarchical Integrity Management for Complex Trusted Platforms

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6.1 Introduction

Trusted Computing has been proposed as a means of providing verifiable trust in a computing platform. However, as virtualization becomes more popular and platform changes (such as security patches) occur more frequently, the established model for Trusted Computing is insufficient to cope in real-world scenarios. We therefore introduce an extensible integrity management framework that is better suited to deal with complicated trust dependencies and change management.

The goal of Trusted Computing is to enable third parties to remotely attest and verify the configuration of a computing platform in a secure manner. Existing *trusted platforms* typically contain a component that is at least logically protected from subversion. The implicitly trusted components of a trusted platform – in particular, the hardware Trusted Platform Module (TPM) – can be used to store integrity measurements, and subsequently report these to users (or remote entities) with a cryptographic guarantee of their veracity. Users can then compare the reported measurements with known or expected values, and thereby infer whether the platform is operating as expected (e.g., it is running the expected software with the expected configuration while enforcing the expected policies).

Present implementations of Trusted Computing technology can take immutable snapshots of a whole platform, which can then be used as proof of trustworthiness [104, 39, 54, 35]. They do not, however, provide more granular verifications of platform components such as individual virtual machines (VMs) and applications. The platform is treated as a whole, and while it is possible to store integrity measurements of VMs and applications, the limited amount of storage in a TPM means that it is not

possible to represent individual components and the dependencies between them. Furthermore, it is not possible to manage changes to measured components. The current scheme advocated by the Trusted Computing Group (TCG) deems all such changes to be malicious [114]. This is certainly impractical for modern server environments, which undergo a constant bombardment of security patches and policy changes. In 2007 alone, Microsoft released 11 security related patches for the Windows operating system [3], while a typical enterprise anti-virus application will undergo two to five updates in an average week [80].

In this paper, we introduce an extensible integrity management framework that addresses these two shortcomings. To improve integrity management, we explicitly represent integrity dependencies between platform components by giving individual registers to each component to store their integrity measurements, and chaining these components together in a dependency graph. To improve change management, we introduce a new distinction between reversible and irreversible changes to measured components. A reversible change is one that can be undone and is guaranteed not to have any permanent effects. The introduction of reversible changes allows the platform integrity to be modified temporarily, for example when a device is hot-plugged and then removed. Although the platform may no longer be considered trustworthy during the time that the change holds, its integrity can be safely restored after the change is undone.

Our resulting framework gives a better understanding of a platform's security properties, which can be used in policy verification. Like existing Trusted Computing implementations, our services can be used to grant access to protected resources (such as encrypted storage) only when the policy is satisfied; however, unlike existing implementations, these policies can be more fine-grained, dynamic, and flexible. Our prototype implementation, built on the Xen virtual machine monitor [30], includes the integrity management framework and a credential manager service, which demonstrates the use of enhanced policy checks to control access to security credentials.

Chapter Outline Section 6.2 outlines the motivation and high-level design for our integrity management framework. Section 6.3 presents the basic framework, which provides integrity services to individual components; Section 6.4 extends this into reversible integrity changes and an explicit dependency graph, and provides use cases for this model. Section 6.5 presents some examples of security services that could make use of our framework. Section 6.6 describes our prototype implementation of the framework and the credential management service on Xen. Finally, in Section 6.7 we draw some conclusions.

6.2 Design Overview

The typical design for a trusted platform comprises a hardware TPM and software integrity management services. These services measure platform components, store integrity measurements as immutable logs and attest these measurements to third parties. The services use the TPM to provide a link with the CRTM. In a non-virtualized platform, with relatively few components to be measured, this model is sufficient. However, it does not scale to complex virtualized platforms that have a plethora of components and dependencies between these components. In this section, we first discuss

the limitations of the existing model. We then present the high-level design goals that motivate our integrity management framework.

6.2.1 Hardware Limitations

Current integrity management systems typically employ the TPM as the sole repository for integrity measurements (see Section 4). Unfortunately, such schemes are fundamentally limited by the hardware capabilities of a TPM:

1. A TPM contains a small, limited amount of memory (PCRs). The TCG specification recommends that a TPM has at least 16 PCRs [114]. Therefore, for portability, we cannot assume that a TPM will have any more than 16 PCRs. Hence, it is not feasible to store individual measurements for a large number of virtualized platform components.

2. The limited number of PCRs is typically addressed by aggregating measurements in the same register. Where two components are independent, this introduces a false dependency between them. Furthermore, the definition of the `extend` function introduces an artificial dependency on the order in which they are aggregated.

3. It is not possible to reverse the inclusion of a measurement in a TPM register. Therefore, it is impossible for a platform component to report a change to its integrity (e.g. by the dynamic loading of some code, or the connection of a new device) and revert back (after unloading/disconnection).

To illustrate these limitations, consider the following example. A server platform hosts tens of small VMs, each of which runs a particular service. To keep track of the platform integrity on a traditional TPM-based system, the measurements must be aggregated, because there are more VMs than PCRs. For example, it might be necessary to store measurements for a virtual network switch and a virtual storage manager in the same PCR, which creates a false integrity dependency between these two VMs. If a malicious change is made to the virtual network switch, and this change is reported to the appropriate PCR, the integrity of the storage manager also appears to be compromised. The same is true for all other VMs whose measurements are aggregated in that PCR.

It would be possible to extend the set of PCRs by giving a virtual TPM to each platform component [13]. However, by allocating independent virtual PCRs to each component, it is no longer possible to represent real dependencies between components[1]. Furthermore, since the virtual TPMs emulate the behavior of a hardware TPM, it remains impossible to revert changes.

6.2.2 High-level Design

It is clear that software measurement support is required to address the limitations of hardware capabilities. We refer to the set of software components that comprise the integrity framework as the *software root of trust for measurement (SRTM)*. These components are part of the platform TCB, and should be isolated from other components;

[1] Some virtual TPM designs share a fixed number of PCRs between all virtual TPMs and the hardware TPM, and these could be used to express dependencies. However, the reliance on the hardware TPM leads to the same limitations as a single-TPM scheme.

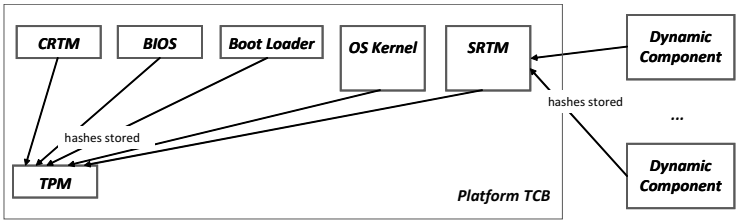


Figure 6.1: The position of the SRTM within the overall integrity management framework.

for example, by virtualization. Dynamic components outside the platform TCB rely on the SRTM to store measurements on their behalf, rather than the underlying TPM. Figure 6.1 illustrates the position of the SRTM within the overall integrity management framework.

Our framework has the following design objectives:

Unlimited measurement storage The framework should allow the storage of individual integrity measurements for an arbitrary number of components.

Explicit dependency representation The framework should allow the explicit and unambiguous representation of an arbitrary number of dependencies between platform components. There should be no false or artificial dependencies introduced by aggregation.

Static integrity management The framework should provide a superset of the functionality of a traditional TPM, with respect to static integrity.

Dynamic integrity management The framework should enable the integrity state of a platform component to revert to a previous trusted state in a controlled and verifiable manner.

Link to hardware TPM The software framework should be linked in a chain of trust to the hardware TPM. This can be achieved by storing the measurements for the SRTM and other static components in the platform TCB (such as the hypervisor and any physical device drivers) in the TPM. As this set of components is small and non-changing, the limitations of a hardware TPM do not come into effect.

Minimal TCB In order to improve the trustworthiness of the framework, the SRTM and other components in the TCB should have a minimal amount of code and size of interface. This paper does not focus on minimizing the TCB, but a possible approach would involve using disaggregation [83].

Platform independence The framework should not be limited to a single hypervisor technology. Although the implementation (see Section 6.6) was carried out using Xen, it should be possible to use alternative technologies, such as VMware [109] or an L4 microkernel [74].

6.3 Basic Integrity Management

In this section, we present a basic design for the SRTM service that we introduced earlier. This platform-independent service provides the minimal functionality needed

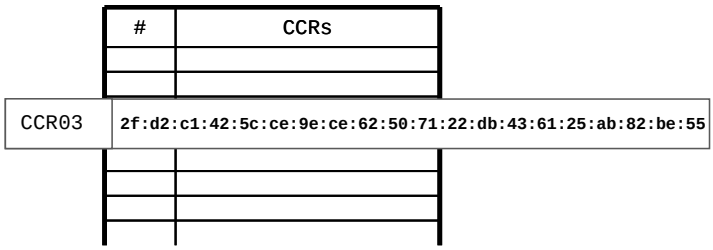


Figure 6.2: Basic integrity management components – Component configuration register table.

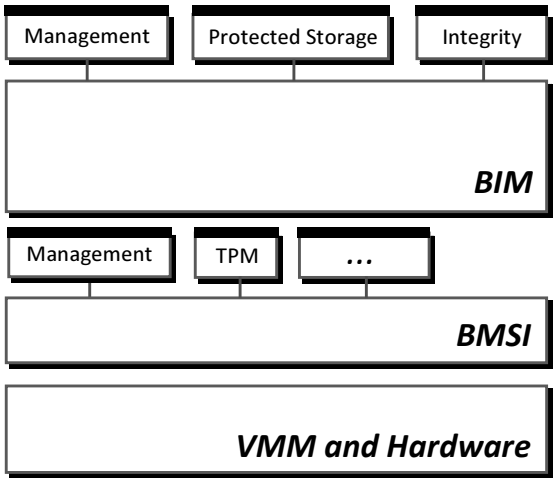


Figure 6.3: Basic integrity management components – The BIM architecture.

to manage the integrity of dynamic (non-TCB) platform components, which will be extended further in Section 6.4. Section 6.3.1 sets out the basic measurement model, while Section 6.3.2 describes the corresponding service architecture and interfaces.

6.3.1 Measurement Model

The Basic Integrity Management (BIM) service stores static integrity measurements of dynamic components that are arranged in a flat hierarchy, such as the one shown in Figure 6.4. Each component has a single Component Configuration Register (CCR) associated with it. A CCR is analogous to a PCR and holds integrity measurements for that component. The measurements are held together in a global CCR table similar to the one depicted in Figure 6.2.

Static Measurements

The BIM measurement model mimics TPM measurement capabilities but stores integrity measurements in software rather than hardware. Each registered dynamic component is assigned a BIM CCR to which its measurements are reported. This is achieved by an `extend` operation, which stores a new measurement in a CCR by hashing it together with the current value of the CCR. Dynamic components use this

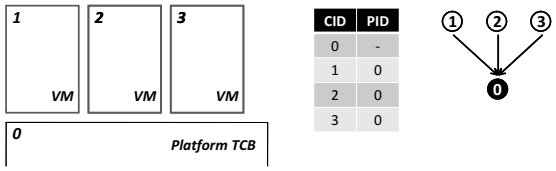


Figure 6.4: Simple integrity use case – a flat hierarchy.

operation to report ongoing measurements when their contents change. For example, a firewall service would extend its CCR if its rule-set was about to be changed. The specifics of when/how measurements are taken is component-dependent, but the logic that performs this activity must be trusted to report changes faithfully. This behavior is assured by the initial measurement of the component by the component that starts it. In the BIM model, this can only be a static (platform TCB) component.

This measurement model provides better scalability than models that use the TPM as the sole repository for measurements. By using software registers, the BIM can store a virtually unlimited number of individual measurements. Hence, no aggregation is needed. However, the measurements are still accumulated and the CCRs are irreversible. That is, recording a measurement M_1, followed by a changed measurement M_2, followed by M_1 again, results in a different value than the original recording of M_1 alone. Hence, components are not allowed to change in any way without permanent loss of integrity. Even if a change is later undone, the component cannot return to its previous trust state. In Section 6.4, we will address this problem by employing dynamic registers for reversible measurements.

Simple Trust Dependency

The BIM service implements a flat hierarchy to capture the integrity dependencies between platform components. In this model, the integrity of dynamic components solely depends on the integrity of the underlying platform TCB. We show an example flat hierarchy in Figure 6.4. The components labeled *one*, *two*, and *three* are virtual machines running directly on the trusted platform. Component *zero* is the platform TCB that includes the SRTM (in this case, the BIM service). Each VM depends only on the platform TCB underneath. If the integrity of the TCB (component *zero*) is compromised, then the integrity of all of the VMs is compromised as well. However, the VMs are independent of one another and therefore do not have a trust dependency. As an example, if the integrity of VM_1 is compromised, the integrity of VM_2 and VM_3 remains intact.

In what follows, we depict the integrity relationships between components using a dependency graph, and represent it using a dependency table. Figure 6.4 shows a simple graph and its dependency table equivalent. For example, the second row in the dependency table states that the integrity of the child component *one* (VM_1) depends on the integrity of the parent component *zero* (TCB).

In the simple BIM model, there is always a single trusted component (the platform TCB) on which all other components depend. This yields the "flat hierarchy" dependency graph and table in Figure 6.4. The flat hierarchy arises, because a dynamic component (such as a VM) can only be started by a trusted component. Since the TCB is static and platform-wide, it is not possible for a dynamic component to start – and hence become a parent of – another dynamic component. Therefore the BIM cannot

Integrity	Description
`extend`	Takes a hash value as an argument and irreversibly extends the component CCR with that hash.
`quote`	Takes arbitrary external data (i.e., nonce) and returns a quotation of the current TCB measurements, the nonce, and the component CCR value signed by a TPM attestation identity key (AIK).
Protected Storage	**Description**
`seal`	Takes data to be protected, seals it to the TPM binding it to current TCB and CCR measurements, and returns the sealed (encrypted) blob.
`unseal`	Takes the sealed blob and unseals and returns the data iff the integrity of the TCB and the component are verified as intact.
Management	**Description**
`register`	Takes the initial measurement, adds the component to the dependency table, and fills the CCR with the initial measurement.
`delete`	Deletes the component and all its sealed data.

Table 6.1: BIM integrity, protected storage, and management interfaces.

manage, for example, the integrity of an application started within a VM. However, the BIM serves as a basis to build the hierarchical model which addresses this limitation, which is introduced in Section 6.4.

6.3.2 The BIM Architecture

As shown in Figure 6.3, BIM services are grouped under three interfaces:

Integrity interface This interface provides functions to report and quote integrity measurements of dynamic (i.e., non-TCB) components. Components use this interface to extend their register values when they detect significant changes to their measured content. A component is only allowed to alter its own register, while an integrity quote can be requested by any entity. Using the underlying TPM interface, the latter operation returns a signed integrity digest that contains the measurements of the dynamic component and the platform TCB. Using this digest, a third party can verify the complete integrity chain.

Protected storage interface This interface provides functions to store and reveal secrets on behalf of dynamic components. These secrets are bound to the integrity of the TCB and the owner component, i.e., they are revealed if and only if the integrity of the component and its ancestors (in the BIM case, the platform TCB) is intact. The BIM uses the underlying TPM interface for sealing and unsealing data to and from the TPM, which automatically implies a verification check on the TCB. Verification of the component's integrity can be done either by the BIM or delegated to a third-party verifier. Our prototype implements the former case in which the BIM needs to store the expected measurements for comparison. We use the TPM sealing operation itself to do so and use the CCR values at the time of sealing as the future expected values. We concatenate these values to the secret and seal the whole blob. The unsealing operation at a later time returns not only the sealed secret but also the expected set of measurements that we compare to the CCR values at that time.

Management interface This interface provides functions to register dynamic components to the framework so that their integrity can be tracked by the BIM. The BIM is a passive service, and so only registered components are tracked. As discussed in the previous sections, the initial measurement of the component is provided from outside by a trusted component that measures and initiates the component. The interface also allows the deletion of components and their sealed data.

Table 6.1 details the individual functions provided by each interface. As shown in Figure 6.3, the BIM, in turn, makes use of the Basic Management and Security Interface (BMSI), which provides a platform-agnostic interface to the underlying hypervisor and hardware TPM. In particular, the BMSI provides functions that enable the BIM to access the TPM and establish a link to the hardware root of trust. The implementation of the BMSI is discussed in more detail in Section 6.6.

6.4 Hierarchical Integrity Management

In this section, we present an enhanced design for the SRTM service that we introduced in Section 6.2. This platform-independent service features dynamic measurements and a component hierarchy that we use to manage the integrity of dynamic (non-TCB) platform components more effectively. We describe the security model for measurements in Section 6.4.1. We describe the service architecture and interfaces in Section 6.4.2.

6.4.1 Measurement Model

The Hierarchical Integrity Management (HIM) service stores integrity measurements in a CCR table as illustrated in Figure 6.2. To overcome the shortcomings of the BIM model (e.g., irreversible measurements), we have extended it by introducing two new concepts: dynamic measurements and hierarchical trust.

Dynamic Measurements

The HIM measurement model enhances the BIM model in two ways. First, HIM allows multiple registers to be assigned to a single dynamic component. This way, component measurements can be tracked with better granularity. Second, HIM supports dynamic measurements that can be reported to a resettable register. This increases flexibility and allows a component to revert back to a trustworthy configuration if permitted by its change policy.

Change types. We distinguish two types of component changes. More specifically: An *irreversible change* is one that requires the component to be restarted before its integrity can be re-established. Such a change is one made to the integrity-critical part of the component; that is, to the code or other data of the component that has a potential impact on the future ability of the component to implement its intended functionality correctly. An example of an irreversible change is a kernel loading an untrusted device driver as the driver may make a change to kernel memory that will persist even after it is unloaded.

A *reversible change* is one in which the component is permitted to re-establish integrity without being completely reinitialized. Such a change is one made to a non-critical part of the component; that is, to code or other data of the component that has no direct or potential impact on the component's future security. A component still

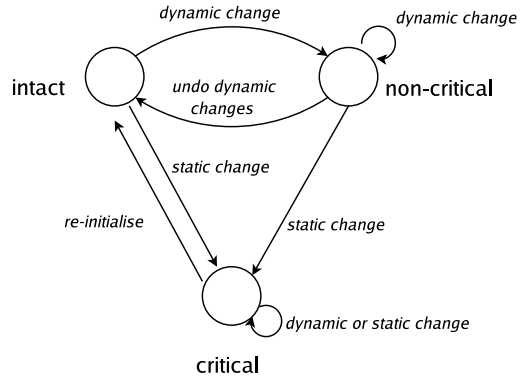


Figure 6.5: Transition diagram for component integrity states. A component in the non-critical state can be made intact by undoing dynamic changes, but the critical state can only return to the intact state by re-initialization.

loses its integrity if a change is made to it. However, depending on the exact nature of the change, we may permit the component to regain integrity (and therefore trust) by undoing the change and returning to its previous state. For example, changes to configuration parameters are often reversible – e.g. changing the identity certificate that a component uses. The integrity management system will need to note such a change in order to fully report the state of the platform, but the certificate may be safely changed back without causing security implications. Another example might be loading a trusted kernel module that is known not to leave any side effects after being unloaded.

The categorization of a change as reversible or irreversible is component-dependent and will be set by each component's own change-type policy. For example, a policy stating that all changes are irreversible reduces to the static measurement model. A component that permits reversible changes is referred to as a *dynamic component* ("dynamic" because its integrity state may change multiple times).

Measurement reporting. Recording dynamic measurements requires two measurement registers, a *static register* and a *dynamic register*, rather than the single register used in the static measurement model. Irreversible changes are reported to the static register in the same way as in the static measurement model; that is, the `extend` operation is used to combine the new measurement with the existing register value to obtain the new register value.

$$extend(R, M) = hash(R||M)$$

where R is the value of the register and M is the measurement.

By contrast, reversible changes are reported to the dynamic register by *replacing* the previous value held in that register, using the `reset` operation.

$$reset(R, M) = M$$

We can see that attempting to reverse an irreversible change does not return the static register to its initial state:

$$R_{final} = extend(extend(R_{initial}, M_2), M_1) = hash(hash(R_{initial}||M_2)||M_1) \neq R_{initial}$$

However, reversing a reversible change *does* return the dynamic register to its initial state:

$$R_{final} = reset(reset(R_{initial}, M_2), M_1) = reset(M_2, M_1) = M_1 = R_{initial}$$

The exact nature of the reporting activity and the corresponding change-type policy is component-dependent. However, the logic that performs this activity must be a part of the initial measurements so that we can trust the component to report the changes to the correct register.

Integrity states. Depending on the measurement values stored in its static and dynamic registers, a dynamic component can be in one of three local integrity states: intact, non-critical, and critical. The component is in the *intact state* if and only if the values in the static and dynamic registers are consistent with the expected measurement values. The component is in the *non-critical state* if and only if the value in the static register is consistent with the expected measurement value but the value in the dynamic register is not. In all other cases, the component is in the *critical state*. As shown in Figure 6.5, the foregoing arrangement enables a dynamic component that has only been subject to non-critical changes to be restored to the intact state. A component that is in the critical state cannot be restored to any other state unless re-initiated with an expected configuration (during which both registers are reset).

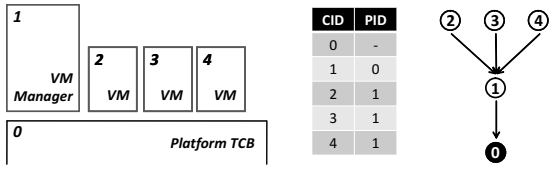
Security states. Depending on the integrity state, a component can be in three security states: trustworthy, secure, and insecure. A component is *trustworthy* if and only if it is in intact state. A component is *secure* if and only if it is in intact or non-critical states. In all other cases, the component is deemed *insecure*.

Example use case for dynamic registers. Digital Rights Management (DRM) services control the distribution of media content onto computing platforms. It is possible that a DRM service will not push video content to a computing accessory if, for example, an external recording device is plugged to it. In this case, software that detects and installs the plug-and-play drivers for the recording device must be part of the static measurements. However, the state in which a recording device is detected in the system can be reported dynamically. In fact, this can be reflected in the dynamic register for a secure DRM player application. As long as the recording device is connected, no content is downloaded. Once the user unplugs the device, the dynamic register is reset and content can be pushed to the player without requiring the application to be restarted.

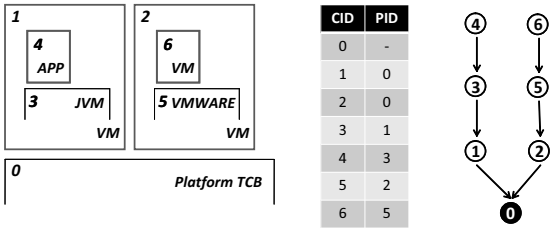
Hierarchical Trust Dependency

We enhance the BIM dependency model by introducing a hierarchy of trust dependencies that we represent as a directed acyclic graph. In such a graph, the edges indicate trust dependencies where the integrity of the component at the origin depends on the integrity of the component at the destination. If the integrity of the destination component is compromised, then the integrity of the origin component is always compromised as well. However, the reverse is not true. To illustrate these more complex trust relationships, consider the following use cases.

In Figure 6.4, we see the simple flat hierarchy as previously described in Section 6.3. The components labeled *one*, *two*, and *three* are virtual machines running directly on the trusted platform. Component *zero* is the platform TCB that includes the SRTM (in this case, the HIM service). Each VM depends only on the platform TCB



(a) Multi-level dependency.



(b) Nested components.

Figure 6.6: Hierarchical integrity use cases.

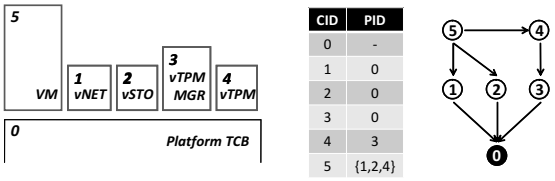
underneath. If the integrity of the TCB (component *zero*) is compromised, then the integrity of all of the VMs is compromised as well. However, the VMs are independent of one another and therefore do not have a trust dependency. As an example, if the integrity of $VM_1$is compromised, the integrity of VM_2 and VM_3 remains intact.

Figure 6.6(a) shows a more complex multi-level dependency. Component *one* is a service that manages the life-cycle of components *two*, *three*, and *four*. All components are virtual machines. The latter VMs are independent of one another, as before, but their integrity depends on that of the domain manager, whose integrity in turn depends on the TCB.

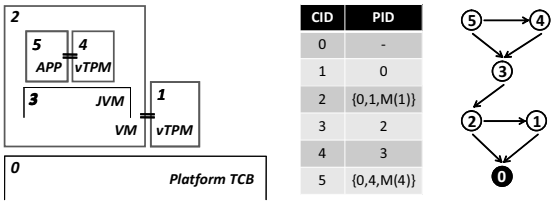
In Figure 6.6(b), we see a nested dependency relationship. Components *one* and *two* are virtual machines, which themselves contain further virtual machines: component *three*, which is a Java virtual machine, and component *five*, which is a VMware hypervisor. These nested virtual machines support guest components: component *four*, a Java application, and component *six*, a VMware guest. Within component *one*, a traditional linear chain-of-trust applies: Java application depends on Java virtual machine depends on operating system. A similar chain can be found within the VMware component. However, these two chains of trust are independent of one another, and both depend ultimately on the underlying platform TCB.

Figure 6.7 illustrates more complicated use cases. In Figure 6.7(a), we see a multiple dependency relationship. Component *five* is a virtual machine that uses services from components *one*, *two*, and *four*. These components are small virtual machines that provide virtual networking, virtual storage, and virtual TPM services, respectively. Further, the integrity of the virtual TPM depends on the integrity of the virtual TPM manager domain (component *three*).

Figure 6.7(b) shows a similar VM grouping example which we intend to explore further in future work. In this example, we use miniature virtual TPM services to assist and enhance the integrity measurement capabilities of the framework. In this design we bind a single virtual TPM to a component (application or VM) and delegate component measurements to this virtual TPM. The virtual TPM then replaces the component CCRs



(a) Disaggregated services.



(b) Virtual TPM binding.

Figure 6.7: More complicated use cases. Dashed lines denote implicit dependency.

to provide more granular run-time measurements for the component it is attached to. The measurements for the virtual TPM service itself is still held by its own CCRs. As an example, the integrity of component *two* now depends on the integrity of component *one* (its attached virtual TPM) and the run-time measurements taken by this virtual TPM (e.g., during authenticated VM_2 bootstrap). We refer to this measurement set as M(*one*). The same holds for the application component *five* and its attached virtual TPM service component *four*. The present HIM implementation does not yet support virtual TPM attachment.

6.4.2 The HIM Architecture

The HIM service implements the same integrity, protected storage, and management interfaces as the BIM service as presented in Section 6.3, but with the following enhancements.

The HIM integrity interface provides an `extend` function that alters the value of the static CCR in the same way as the BIM equivalent. To support dynamic measurements, the interface also provides a `reset` function that is used to report to the dynamic register and overwrite its value. In addition, to support hierarchical integrity dependency, the `quote` function is modified. This function now returns the aggregated integrity measurements of the component in question. Specifically, the signed quote now contains the TCB integrity measurements plus the measurements of the component and all its ancestors hashed in a single value.

In the HIM protected storage interface, the `seal` and `unseal` functions are enhanced to support component dependency and dynamic measurements. The `seal` function now binds the stored secret to the integrity of all the trust chains that reach the component in question from the TCB; that is, the subgraph of all paths from that component to the root TCB. Hence the integrity state of components not on a path between that component and the TCB is ignored. For example, in the nested use case in Figure 6.6(b), an integrity compromise in the VMware compartment will not affect the ability of the Java application to unseal previously sealed information, as long as the

Java compartment remains intact.

Lastly, the HIM management interface provides `register` and `delete` functions. The `delete` function is the same as in the BIM. However, the `register` function now takes a dependency list as a parameter that specifies additional ancestor components the component depends on besides the one that registers the component.

6.5 Policy Verification for Security Services

In this section, we introduce example security services that leverage the HIM framework for policy verification and access control. Our examples include a credential management service (Section 6.5.1), a virtual TPM service (Section 6.5.2), and a virtual network service (Section 6.5.3).

6.5.1 Credential Management Service

Protected storage services provide secure access to secrets that are sealed to the underlying TPM on behalf of their owners. It is expected that these services retain control over these secrets and enforce the associated access control policies at all times. By contrast, most storage services such as [104] and the HIM provide one-time verification, and are therefore susceptible to a time-of-check to time-of-use vulnerability. This occurs because these services release the stored secret to the requesting component once they verify the necessary policies (e.g., HIM unseal successfully verifies the aggregate integrity). Once the secret is revealed, these services can no longer restrict access to it if the component undergoes a malicious change.

To enable ongoing policy verification and enforcement, we designed and implemented a credential management service (CMS) that uses the integrity management framework to provide secure access to secrets while maintaining control at all times. Unlike the HIM unseal operation, CMS credentials are never revealed to requesting services directly but are always held securely by the CMS. In essence, the CMS is a reference monitor that mediates and provides access to secured data through a well-defined interface.

The CMS interface is comprised of management and service interfaces. Components use the management interface to register component credentials with the CMS. To do so, the `register` function takes the credential as input and seals it to the underlying TPM. The interface also provides a `discard` function which deletes the stored credential. The service interface provides access to the credentials through a generic `access` function. We have designed this interface as an extensible plug-in interface; that is, the exact nature of the interface depends on the nature of the stored credential and the type of functionality needed. For example, if the stored credentials are cryptographic keys, we offer a plug-in service that provides encryption/decryption capabilities so that components can use the interface to encrypt/decrypt data without seeing the actual key. Regardless of the functionality provided, the CMS uses the HIM to verify the aggregate integrity prior to each access to the secret.

6.5.2 Virtual TPM Service

A natural extension to the CMS functionality would be to provide a miniature TPM interface to the various platform components, as illustrated in Figure 6.7(b). This enables these components to have a standardized interface as in [114] to prove their integrity

and provides a strong identity for each component. Such an approach has already been taken through TPM virtualization [13] which gives each VM a TPM interface implemented by a virtual TPM service. However, it is not yet clear what the best mechanism is for establishing a secure binding between a virtual TPM and its platform TCB.

Our framework could be used to bridge the gap between virtual TPM services and the platform TCB. For example, a central trusted CMS service could be used as the single secure repository for virtual TPM keys. Access to these keys would require verification of the complete HIM integrity chain, including verification of the platform TCB. For example, to sign a quote request, a virtual TPM would use the CMS interface to gain access to its signing key.

6.5.3 Virtual Network Service

Virtualization provides direct isolation of computing resources such as memory and CPU between guest operating systems on a physical platform. However, the network remains a shared resource as all traffic from guests will eventually end up on the same physical medium. Various mechanisms can be used to provide network isolation between network domains, as described in [19]. In general, encryption must be used for isolation when network traffic is delivered over an untrusted shared physical medium.

Using our framework in combination with the CMS, one could design a virtual network (vNET) service which provides isolation through an encryption layer such as IPSEC. In this setting, the vNET service would store its credentials (e.g., network encryption key) in the CMS, in combination with the expected CCR values of the service and any ancestor service it depends on (including any potential network configuration information). Because the key is held by the CMS and not revealed to the vNET service, any change in the integrity of the service or its ancestor components would result in the network link becoming unavailable for the VM connected to this specific vNET. As a result, the capability of a VM to communicate with its peer within a considered domain would implicitly prove its trustworthiness, which would provide continuous authentication as opposed to relying only on an initial handshake as most network authentication mechanisms do.

6.6 Implementation in Xen

In this section, we describe a prototype implementation of the integrity management framework and the credential management service on the Xen virtual machine monitor [30]. The implementation features the management and service interfaces of both. Note that although we present our implementation with Xen, the framework could equally be implemented on an alternative virtualized or microkernel-based platform (e.g., the L4/Fiasco [70] microkernel).

6.6.1 Infrastructure Overview

The various components of the integrity management framework are provided by one or more virtual machines, running on top of the Xen virtual machine monitor. The use of virtualization isolates the trusted platform from a misbehaving guest operating system, and all communication with the trusted platform passes through well-defined interfaces. Our implementation is based on Xen version 3.0.4, a VMM for the IA32

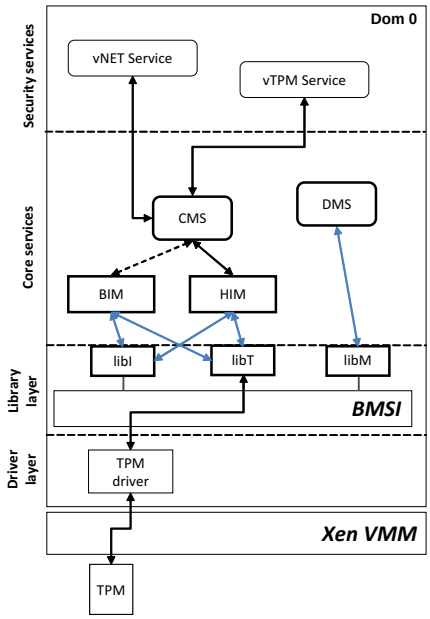


Figure 6.8: Illustration of the prototype in a layered stack.

platform, with the VMs running a paravirtualized version of Linux 2.6.18. For inter-domain communication, we employ the light-weight communication library introduced in [8].

Figure 6.8 illustrates our implementation on Xen. In the present prototype, all framework components and the CMS are implemented as libraries and services running in the Xen privileged management domain Dom0. However, as we have defined interfaces between each of the components, it should be straightforward to move towards a disaggregated approach as described in [83]. The framework components are arranged in a layered stack. At the lowest layer is the basic management and security interface (BMSI) that provides libraries for domain life-cycle management (libM), basic TPM access (libT), and integrity management (libI). At the core services layer are the integrity manager services BIM and HIM that provide basic and hierarchical integrity management, respectively. Also in this layer are the CMS and the domain management service (DMS). At the highest layer are the security services that use the framework for various purposes. The platform TCB consists of the static components up to and including the SRTM (the BMSI libraries and the integrity managers). However, for simplicity, we also include the CMS in the platform TCB. The measurements of these components are reported to the underlying TPM. The application TCB consists of the platform TCB plus the security services that run on top of it. The measurements of the latter are reported to the SRTM.

6.6.2 Component Design

In the present prototype, we have implemented the highlighted components depicted in Figure 6.8, namely the BMSI libraries, BIM and HIM services, CMS, and DMS. In this section, we present the details of these components including both the BIM and HIM; however, due to space constraints, we present an example use case that uses only

the HIM.

BMSI Libraries

The Basic Management and Security Interface (BMSI) provides a common and extensible interface to the underlying hypervisor (i.e., Xen) and the TPM. The BMSI provides libraries for domain life-cycle management (libM), basic TPM access (libT), and integrity management (libI).

libM This library provides hypervisor-agnostic management functions to upper layers. At its lowest level, the library manages allocatable resources called *Protection Domains (PDs)*. A PD is an executable component that receives an allocation of memory and CPU cycles, and is scheduled by the hypervisor. On Xen platforms, a PD is equivalent to a Xen domain (virtual machine). In this prototype, we use libM to implement the Domain Management Service (DMS). This service manages the life-cycle of PDs and uses the integrity managers to keep track of PD integrity. We refer the reader to [83] for further details on the libM and DMS implementation.

libT This library provides the minimal functionality to access the integrity and protected storage interfaces of the TPM. Security services (e.g., BIM and HIM) use this library to obtain a signed quotation of the TCB measurements and to seal/unseal data to/from the TPM. To do so, libT uses the TPM functions `TPM_Quote()`, `TPM_Seal()`, and `TPM_Unseal()` as described by the TPM specification [114].

libI This library stores and provides access to the integrity measurement and dependency tables. The `getMeasurement()` function returns a measurement list that includes the integrity measurements of the component and its ancestors. In the BIM case, a single value is returned. The `setMeasurement()` function extends the value of the component register. The `resetMeasurement()` function overwrites the value of the dynamic register. The `addComponent()` function adds an entry to the dependency table and sets its dependencies as specified. It also adds an entry to the measurements table and records the initial measurements. The `deleteComponent()` function checks that the specified component has no successors and removes it from the table.

Component Interactions

The BIM and HIM services implement the interfaces presented in Sections 6.3 and 6.4, respectively. Similarly, the CMS service implements the interfaces presented in Section 6.5.1 and uses a cryptographic service as a plug-in for block encryption and decryption. On a Xen platform, we use these services to manage the integrity of VMs and applications running on these VMs.

VM integrity management is incorporated into VM life-cycle management. To assist both, the DMS uses the BMSI library libM and the HIM service. The VM start-up phase in Figure 6.9 depicts the interaction among these components. During this phase, the DMS invokes libM, which prepares resources for the VM, measures the VM image (comprising the kernel, an optional initial ramdisk and command-line parameters), and stores the measurement in the CCR for that VM. This performs a function similar to a secure bootloader, and it is the responsibility of the kernel to measure any components

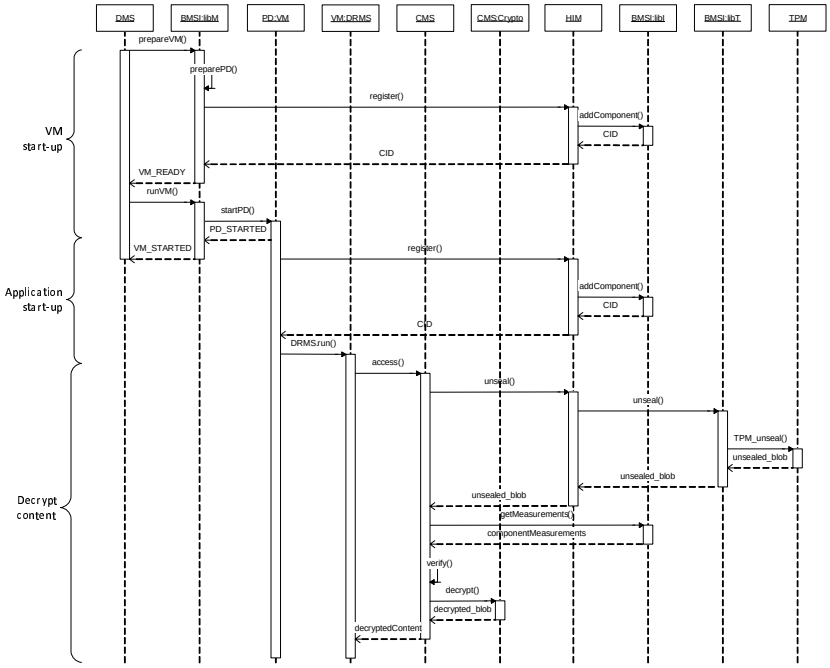


Figure 6.9: Sequence diagram of interactions between the framework services for a DRM application use case.

which it subsequently loads. The DMS also registers the new VM with the HIM service, and configures any dependencies between the new VM and existing VMs. The HIM uses libI to store this information in the measurement and dependency tables. Following the successful completion of the above steps, the DMS starts the VM.

The HIM service additionally allows applications running in VMs to be registered with the framework. The application start-up phase in Figure 6.9 depicts the case in which the VM that was started in the previous phase loads and registers a DRM service with the HIM. In this case, the VM becomes an ancestor of the service and provides its initial measurements. As a result, the cumulative integrity of the service now includes the VM's measurements as well as the platform TCB measurements.

The last phase in Figure 6.9 depicts a use case in which the DRM service that was started in the previous phase attempts to decrypt encrypted media content using a key that is stored on the TPM on behalf of this service. The DRM service invokes the CMS service interface to request access to this key. The CMS then invokes HIM unseal to retrieve the key from the TPM. HIM unseals the key if and only if the underlying policies regarding the key's release are satisfied. In this case, the key is unsealed from the TPM and returned to the CMS if the integrity of the platform TCB is intact. On receiving the key from the HIM, the CMS performs further verification. It compares the expected CCR values of the DRM service and its ancestor VM (unsealed along with the key) to the current CCR values. If the measurements match, the CMS uses its cryptographic service to decrypt the block, which is then returned to the DRM service. Note that any subsequent access requests to the key will also follow a similar verification cycle, with the exception that HIM (hence TPM) seal is omitted because the CMS caches the key internally.

6.7 Conclusions

In this paper, we have introduced a novel integrity management framework that improves on the integrity measurement and policy verification capabilities of present Trusted Computing solutions. In particular, our framework is able to cope with the proliferation of measured components and dependencies between them as well as dynamic changes to platform components. In essence, the framework implements a small, software-based root of trust for measurement (SRTM) that provides a secure link to the core root of trust for measurement (CRTM). We have implemented our framework on the Xen virtual machine monitor and proposed several ways in which security services could take advantage of this architecture for policy verification and access control.

We anticipate integrity and trust management to become especially useful for application and service level components. We will therefore continue to investigate further potential uses for our framework by user level applications. In the short term, we plan to implement CMS-aware services such as a virtual network service based on [19] that uses the CMS to store encryption keys. The virtual TPM service is also particularly interesting. In the long term, we plan to investigate various ways of exploiting our framework to help enhance the security properties of virtual TPM services (e.g., their binding to the platform TCB). Conversely, we plan to use virtual TPM services to help enhance the measurement capabilities of the HIM framework and provide more granular run-time measurements compared to a single CCR.

Chapter 7

Towards Automated Provisioning of Secure Virtualized Networks

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7.1 Introduction

Virtualization allows the abstraction of the real hardware configuration of a computer system. A computer uses a layer of software, called the Virtual Machine Monitor (VMM), to provide the illusion of real hardware for multiple virtual machines (VMs). Inside each VM, the operating system (often called the *guest* OS) and applications run on the VM's own virtual resources such as virtual CPU, virtual network card, virtual RAM, and virtual disks. A VMM can be hosted directly on the computer hardware (e.g., Xen [11] or VMware ESX) or within a host operating system (e.g., VMware Player).

A challenge of virtualization is isolation between potentially distrusting virtual machines and their resources that reside on the same physical infrastructure. Machine virtualization alone provides reasonable isolation of computing resources such as memory and CPU between guest domains. However, the network remains to be a shared resource as all traffic from guests eventually pass through a shared network resource (e.g., a physical switch) and end up on the shared physical medium. Today's VMM virtual networking implementations provide simple mechanisms to bridge all VM traffic through the actual physical network card of the physical machine. This level of isolation can be sufficient for individual and small enterprise purposes. However, a large-scale infrastructure (e.g., a virtualized data center) that hosts services belonging to multiple customers require further guarantees on customer separation, e.g., to avoid accidental or malicious data leakage.

Our focus in this paper is security-enhanced network virtualization, which (1) allows groups of related VMs running on separate physical machines to be connected together as though they were on their own separate network fabric, and (2) enforces cross-group security requirements such as isolation, authentication, confidentiality, integrity, and information flow control. The goal is to group related VMs (e.g., VMs belonging to the same customer in a data center) distributed across several physical

machines into *virtual enclave networks*, so that each group of VMs has the same protection as if the VMs were hosted on a separate physical LAN. Our solution for achieving this goal also takes advantage (whenever possible) of the fact that some VMs in a group may be co-hosted on the same hardware; it is not necessary to involve the physical network during information flow between two such VMs.

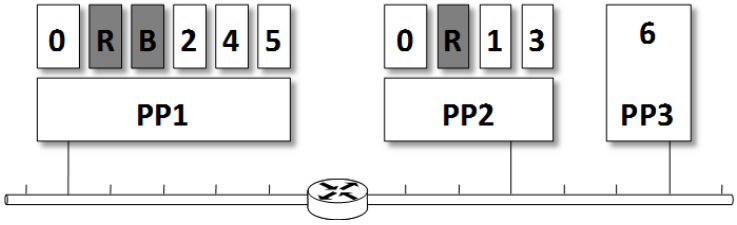
The concept of Trusted Virtual Domains or TVDs was put forth by Bussani et al. [16] to provide quantifiable security and operational management for business and IT services, and to simplify overall containment and trust management in large distributed systems. Informally speaking, TVDs can be thought of as security-enhanced variants of virtualized network zones, in which specified security policies can be automatically enforced. We describe the first practical realization of TVDs using a secure network virtualization framework that guarantees reliable isolation and flow control requirements between TVD boundaries. The requirements are specified by TVD policies, which are enforced dynamically despite changing TVD membership, policies, and properties of the member VMs. The framework is based on existing and well-established network virtualization technologies such as Ethernet encapsulation, VLAN tagging, virtual private networks (VPNs), and network access control (NAC) for configuration validation.

Our main contributions are (1) combining standard network virtualization technologies to realize TVDs, and (2) orchestrating them through a management framework that is geared towards automation. In particular, our solution aims at automatically instantiating and deploying the appropriate security mechanisms and network virtualization technologies based on an input security model, which specifies the required level of isolation and permitted network flows.

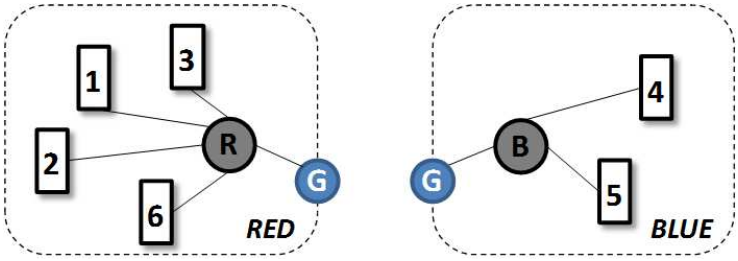
Chapter Outline The remainder of this chapter is organized as follows. In Section 7.2, we provide an overview of our security objectives and describe the high-level framework to achieve the objectives. We introduce the networking components required for our framework in Section 7.3 and describe how they can be orchestrated to enforce TVD policies. In Section 7.4, we present the TVD security model and the components constituting the TVD infrastructure. In Section 7.5, we cover the dynamic aspects of TVD deployment including TVD establishment, population, and admission control. In Section 7.6, we describe a Xen-based prototype implementation of our secure virtual networking framework.

7.2 Design Overview

At a high level, our network virtualization framework comprises virtual networking (VNET) and TVD infrastructure. In essence, the VNET infrastructure abstracts away the underlying network fabric and realizes the mapping from physical to logical network topologies. The TVD infrastructure configures the network devices and instantiates the appropriate set of mechanisms to realize TVD policies in conjunction with the VNET infrastructure. Figure 7.1 illustrates one such physical to virtual mapping. Using this abstraction, the logical layout as observed by the TVD owners (Figure 7.1(b)) remains agnostic to the actual layout of the physical infrastructure (Figure 7.1(a)). In this section, we provide an overview of the mechanisms that enable this mapping while enforcing TVD policies within and across customer domains.



(a) Physical view.



(b) Logical view.

Figure 7.1: An example TVD-based virtual network mapping: R and B are virtual switches for *red* and *blue* TVDs; G is a virtual gateway that controls inter-TVD communication (not shown in (a) for clarity).

7.2.1 High-level Objectives

This paper mainly considers virtual data centers that host multiple customers with potentially conflicting interests. In this setting, our secure networking framework helps meet the following functional and security objectives:

Virtual Networking The framework should support arbitrary logical network topologies that inter-connect virtualized and non-virtualized platforms alike.

Customer Separation The framework should provide sufficient separation guarantees to customers such that data leakage between customer domains is minimal[1].

Unified Policy Enforcement The framework should enforce domain policies within and across customer domains in a seamless manner. The policies define integrity, confidentiality, and isolation requirements for each domain, and information flow control requirements between any two domains.

Autonomous Management The framework should deploy, configure and manage the networking and security mechanisms in an automated manner.

7.2.2 Overview of the VNET Infrastructure

Our VNET infrastructure inter-connects groups of related VMs as though they were on their own separate network fabric. As a result, pockets of virtual LAN (VLAN) segments are formed across the data center. The central VLAN element is a virtual switch, which regulates and isolates the network traffic for each segment. There is a single virtual switch per VLAN segment. A VM appears on a particular VLAN segment if one of its virtual network interface cards (vNICs) is "plugged" into one of the switch ports on the virtual switch forming that segment.

The virtual switch behaves like a normal physical switch and coordinates traffic to and from member VMs. Figure 7.1(b) illustrates two virtual switches R and B that coordinate the subnet communication for *red* and *blue* segments. The virtual switch logic resides either on the host OS of the virtualized platform, or on a dedicated VM as shown in Figure 7.1(a). To satisfy security properties even when the traffic traverses untrusted communication medium, additional VPN modules may be employed. Similarly, additional network admission control (NAC) modules may be employed to authenticate the VMs prior to VLAN admission.

7.2.3 Overview of the TVD Infrastructure

A TVD is represented by a set of distributed virtual processing elements (VPE) (e.g., VMs and virtual switches) and a communication medium interconnecting the VPEs. The TVD provides a policy and containment boundary around those VPEs. At a high level, the TVD policy has three aspects: (1) membership requirements for the TVD, (2) interaction among VPEs of that TVD, and (3) interaction of the TVD with other TVDs and the outside world. VPEs within each TVD can usually communicate freely and securely with each other. At the same time, they are sufficiently isolated from outside VPEs including those belonging to other TVDs. From a networking perspective, isolation loosely refers to the requirement that a dishonest VPE in one TVD cannot send

[1] Indirect communication channels such as covert and subliminal channels are outside the scope of this paper.

messages to a dishonest VPE in another TVD, unless the inter-TVD policies explicitly allow such data flow.

Each TVD has an associated *infrastructure* whose purpose is to provide a unified level of security to member VPEs, while restricting the interaction with VPEs outside the TVD to pre-specified, well-defined means only. Unified security within a domain is obtained by defining and enforcing *membership and communication requirements* that the VPEs and networks have to satisfy before being admitted to the TVD and for retaining the membership. Unified security across domains is obtained by defining and enforcing *flow requirements* that the VPEs have to satisfy to be able to interact with VPEs in other TVDs. To do so, each TVD defines rules that regulate in-bound and out-bound network traffic and restrict communication with the outside world.

The central element in the TVD infrastructure is the TVD master which keeps track of high-level TVD policies and secrets. TVD masters maintain control over member VPEs using TVD proxies. There is a single TVD master per TVD and a single TVD proxy per physical platform. A VPE appears on a particular TVD if it complies with the membership requirements enforced by the TVD proxy controlling that domain on that physical platform. VPEs communicate freely within a TVD. Communication across TVDs is controlled by entities configured by TVD proxies.

7.2.4 Overview of TVD-based Virtual Networking

VNET and TVD infrastructures play complementary roles in meeting the aforementioned high-level objectives. In particular, the VNET infrastructure provides the tools and mechanisms to enable secure virtual networking in a data center. The TVD infrastructure defines unified domain policies, and automatically configures and manages these mechanisms in conjunction with the policy. It also provides domain customers the necessary tools to manage their TVDs. Any action by the customer is then seamlessly reflected on the underlying VNET infrastructure in an automated manner.

This two-layer architecture has many advantages:

1. The policy (TVD) and enforcement (VNET) layers are clearly separated from each other, which yields a modular framework. As a result, a plethora of VNET technologies can be deployed underneath the TVD infrastructure in place of the current offering.

2. The policy layer is not restricted to network-type policies only. For example, similar policies can be devised to configure and manage mechanisms to share virtual storage between VMs in a single TVD.

3. The policy layer provides a high-level abstraction of the infrastructure to customers, which in turn yields simpler administrative tools that are agnostic to the complexities of the underlying VNET technology.

4. Automated deployment, configuration and management of customer TVDs are desirable in a virtual data center that can effectively reduce operational costs by easing the management burden, and mitigate insider threat by automating administrative tasks.

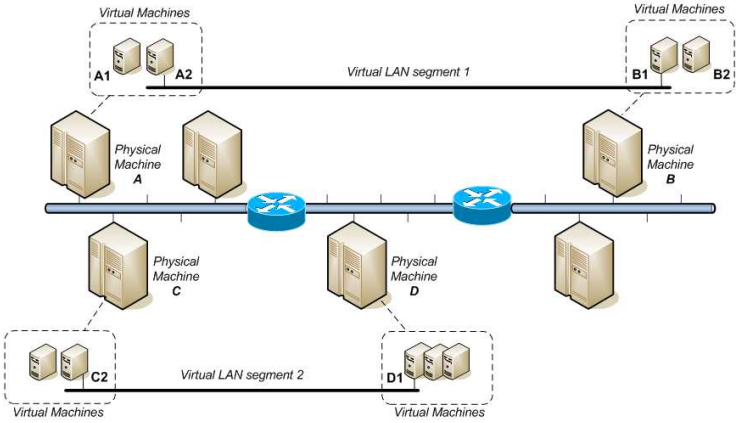


Figure 7.2: An Example Physical to Virtual Mapping with two Virtual LAN Segments.

7.3 Networking Infrastructure

In this section, we present the network virtualization layer that enables the creation of arbitrary virtual topologies on top of the underlying physical network. This layer also provides the basic mechanisms to help enforce TVD admission and flow policies at a higher level.

7.3.1 Aims and Requirements

Functional Requirements

The main aim of the VNET infrastructure is to allow groups of related VMs running on separate physical machines to be connected together as though they were on their own separate network fabric. For example, a group of related VMs may have to be connected directly together on the same VLAN segment even though, in reality, they may be at opposite ends of a WAN link separated by many physical LAN segments. Further, multiple segmented virtual networks may have to be established on a single physical network segment to achieve improved security properties such as network isolation. Third, the VNET infrastructure needs to inter-operate with existing non-virtualized entities (e.g., standard client machines on the Internet) and allow our virtual networks to connect to real networks.

Figure 7.2 illustrates an example physical to virtual mapping. Physical machine A hosts two VMs. One of them (A2) is connected into VLAN segment 1. Physical machine B also hosts two VMs. One of them (B1) is also connected into VLAN segment 1. VM A1 and VM B1 appear as though they are connected together via a single LAN segment even though in reality the physical network connecting them consists of multiple LAN segments interconnected via routers. VMs C2 and D1 are also connected together via VLAN segment 2. Traffic is isolated between VLAN segments so machines C2 and D1 cannot see the traffic passing between A2 and B1 even though they are sharing the underlying physical network infrastructure.

Attacker Model

Our research specifically focuses on virtual data centers that aim to separate customer networks and protect network traffic even when the transmission lines are untrusted. We mainly focus on network separation and protection; however, our policies and infrastructure are generic and can be applied to other requirements such as storage isolation. A particular threat we consider here is the threat of malicious data center users that can try to gain access to other customer networks for the purpose of stealing sensitive information or rendering those systems unusable. Additional threats could include malicious users tapping untrusted transmission lines when data travels across a WAN; malicious users injecting traffic into unauthorized network segments; or in a more extreme case, a compromised virtual network device leaking all data passing through the switch or router. These risks clearly reduce the benefits of machine consolidation in a virtual data center. The risks also limit the guarantees that infrastructure owners would put into service-level agreements when compared to the case of physical isolation of the networks belonging to different customers.

Security Requirements

Our VNET infrastructure adheres to following security requirements to address these threats in a virtual data center:

1. The infrastructure must provide logical isolation of VLAN segments. A rogue segment must not be able to gain control of devices that belong to other VLANs in the virtual data center.

2. The infrastructure must protect and separate the traffic generated by each VLAN segment. VLANs must not be able to see, intercept or inject packets into traffic originating from other VLANs.

3. Network connections can be established between VMs that belong to the same VLAN segment and between VMs that belong to different VLAN segments (i.e., TVDs). The infrastructure must secure these connections whenever they are established over untrusted physical medium.

4. The infrastructure must authenticate and sanity-check all VNET entities (virtual or physical) including the management entities prior to admission to the particular VLAN segment.

7.3.2 Networking Infrastructure

One option for virtual networking is to virtualize at the IP level. However, to also support non-IP protocols and IP support services (such as ARP) that sit directly on top of the Ethernet protocol, we have chosen to virtualize at the Ethernet level.

Architecture Overview

Our VNET infrastructure employs a mixture of physical and virtual networking entities to inter-connect physical and virtual machines with respect to the logical topology. Physical networking entities include standard physical networking devices such as VLAN-enabled switches and routers. Virtual networking entities include virtual network interface cards (vNICs), virtual switches, virtual routers, and virtual gateways.

Table 7.1: Virtual components of the VNET infrastructure.

VNET COMPONENT	FUNCTIONALITY
Virtual NIC	Connects a virtual machine to a virtual switch
Virtual Switch	Provides centralised switching functionality per subnet Performs address mapping and translation using switching table Provides logical traffic isolation between VLAN segments
EtherIP Module	Tunnels Ethernet and 802.3 packets via IP datagrams Expands a LAN over a WAN or MAN
VLAN Module	Tags Layer 2 frames with the corresponding VLAN identifier Provides traffic isolation on the physical switch
VPN Module	Encapsulates traffic in IP packets Provides confidentiality and integrity over insecure medium
NAC Module	Enables port-based access control and compliance checking prior to network admission
Virtual Router	Routes Layer 3 traffic between VMs on different VLAN segments
Virtual Gateway	Advertises routing information about the virtual network behind it Converts packets to and from the VNET encapsulated format

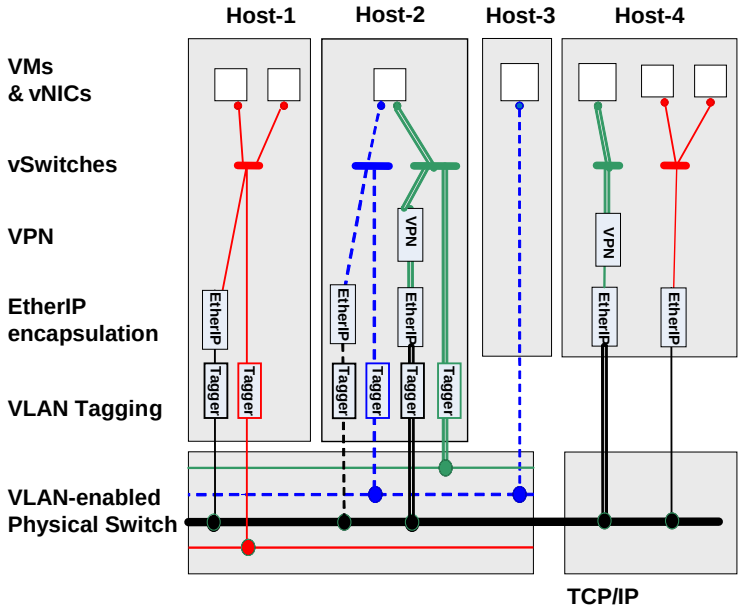


Figure 7.3: Components of the Secure Virtual Networking Infrastructure.

Table 7.3.2 provides a list of all virtual networking devices we employ in our networks alongside standard physical devices (e.g., physical switches). In essence, a central virtual switch per virtual LAN (VLAN) segment uses a combination of these devices and modules to enable the virtualization of the underlying network and secure the communication. We provide details on each virtual networking device in the following sections.

Figure 7.3 illustrates how these components can be composed into a secure VNET infrastructure that provides isolation between different VLANs (i.e., TVDs), where each TVD is represented by a different color (red (solid), green (dashed), or blue (double) line). Abstractly speaking, it is as if our secure VNET framework provides colored networks (in which a different color means a different TVD) with security guarantees (such as confidentiality, integrity, and isolation) to higher layers of the virtual infrastructure. VMs connect to the network using their vNICs through the corresponding virtual switch for that VLAN; whereas a non-virtualized physical host, such as Host-3 in Figure 7.3, is directly connected to a VLAN-enabled physical switch without employing a virtual switch. A VM can be connected to multiple VLAN segments using a different vNIC for each VLAN segment; hence, the VM can be a member of multiple TVDs simultaneously. For example, the lone VM in Host-2 is part of two VLAN segments, each represented by a virtual switch with a different color; hence, the VM is a member of both the blue and green TVDs.

Virtual Switching

Virtual switches enable network virtualization and co-ordinate all communication within and across VLAN segments. Each VM has a number of vNICs where each can be associated with at most one VLAN segment. Each VLAN segment is represented

by a virtual switch or a *vSwitch*, which behaves like a normal physical switch for that segment. A VM appears on a particular VLAN if one of its vNICs is "plugged" into one of the switch ports on the vSwitch. Ethernet broadcast traffic generated by a VM connected to the vSwitch is passed to all VMs connected to that vSwitch. Like a real switch, the vSwitch also builds up a forwarding table based on observed traffic so that non-broadcast Ethernet traffic can be delivered point-to-point to improve bandwidth efficiency.

To enable network virtualization, our vSwitches encapsulate or tag VM Ethernet frames with VLAN identifiers using the EtherIP [52] and IEEE 802.1Q [49] modules, respectively. Ethernet frames originating from the source host are handled differently depending on whether the source host is virtualized and whether the destination host resides in the same LAN. For a virtualized domain (e.g., Host-1 in Figure 7.3), each frame is tagged using a VLAN tagging module. If the destination of the Ethernet frame is a VM on another host that is connected to the same VLAN-capable switch (e.g., another physical domain in a data center), this tag indicates the VLAN segment to which the VM belongs. If the destination is a host that resides outside the LAN domain (e.g., Host-4), the VLAN tag forces the switch to bridge the connection to an outgoing WAN line (indicated by the black (thick) line in the VLAN-enabled physical switch of Figure 7.3) that is connected to a router for further packet routing. In this case, the VM Ethernet frames are encapsulated in IP packets to indicate the VLAN segment membership. Lastly, if a non-virtualized physical host is directly connected to the VLAN switch (e.g., Host-3), no tagging is required for the outgoing connection from the host's domain. In the absence of a trusted physical network, each VLAN segment can employ an optional VPN layer to provide authentication, integrity, and confidentiality.

EtherIP Encapsulation

EtherIP is a standard protocol for tunneling Ethernet and 802.3 packets via IP datagrams and can be employed to expand a LAN over a Wide or Metropolitan Area Network [52]. We employ EtherIP encapsulation as the standard mechanism to insert VLAN membership information into Ethernet/802.3 frames. To do so, each tunnel endpoint uses a special network device provided by the operating system that encapsulates outgoing Ethernet/802.3 packets in new IP packets. We insert VLAN membership information (i.e., the VLAN identifier) into the EtherIP header of each encapsulated packet. The encapsulated packets are then transmitted to the other side of the tunnel where the embedded Ethernet/802.3 packets are extracted and transmitted to the destination host that belongs to the same VLAN segment.

Address Mapping The vSwitch component maps the Ethernet address of the encapsulated Ethernet frame to an appropriate IP address. This way, the encapsulated Ethernet frame can be transmitted over the underlying physical network to physical machines hosting other VMs connected to the same LAN segment that would have seen that Ethernet traffic had the VMs actually been on a real LAN together. The IP address chosen to route the encapsulated Ethernet frames over the underlying physical network depends upon whether the encapsulated Ethernet frame is an Ethernet broadcast frame and also whether the vSwitch has built up a table of the locations of the physical machines hosting other VMs on a particular LAN segment based on observing traffic on that LAN.

IP packets encapsulating broadcast Ethernet frames are given a multicast IP address and sent out over the physical network. Each VLAN segment has an IP multicast address associated with it. All physical machines hosting VMs on a particular VLAN segment are members of the multicast group for that VLAN segment. This mechanism ensures that all VMs on a particular VLAN segment receive all broadcast Ethernet frames from other VMs on that segment. Encapsulated Ethernet frames that contain a directed Ethernet address destination are either flooded to all the VMs on a particular VLAN segment (using the IP multicast address as in the broadcast case) or sent to a specific physical machine IP address. This depends upon whether the vSwitch component on the encapsulating VM has learned the location of the physical machine hosting the VM with the given Ethernet destination address based on traffic observation through the vSwitch.

Requirements Revisited EtherIP can be used over arbitrary Layer 3 networks. The decision to encapsulate Ethernet frames from VMs within IP packets allow us to connect different VMs to the same VLAN segment so long as the physical machines hosting those VMs have some form of IP based connectivity between them, even a WAN link. There are no restrictions on the topology of that physical network.

To allow routing within virtual networks we use virtual routers that reside on VMs with multiple vNICs. The interface cards are plugged into ports on the different vSwitches that it is required to route between. Standard routing software is then configured and run on the routing VM to provide the desired routing services between the connected VLAN segments.

To allow for communication with non-virtualized systems we provide a virtual gateway that simply is another VM with two vNICs. One vNIC is plugged into a port on a vSwitch; the other one is bridged directly on to the physical network. The gateway has two main roles: (1) It advertises routing information about the virtual network behind it so that hosts in the non-virtualized world can locate the VMs that reside on a virtual network, and (2) it converts packets to and from the encapsulated format required of our virtual networks.

VLAN Tagging

VLAN tagging is a well-established VNET technology that provides isolation of VLAN segments on physical network equipment [49]. We employ VLAN tagging as an alternative to Ethernet encapsulation for efficiency purposes. For example, in a virtualized datacenter a VLAN-enabled switch may be used that yield increased performance over EtherIP encapsulation.

Each VLAN segment employs its own VLAN tagging module to tag its Ethernet frames. This module resides within the hypervisor or host OS that facilitates the networking capabilities, captures packets originating from VMs, and tags those with the ID of the VM's VLAN before sending them onto the physical wire. On the receiving side, the module removes the VLAN tag and passes the packets untagged into the destination VM(s). Packets are only tagged when they have to be transmitted over the physical network. VMs are unaware of the VLAN tagging and send/receive packets without any VLAN information.

To handle VLAN tagged packets, the physical network equipment needs to support IEEE 802.1Q and be configured accordingly. As an example, if a machine hosts a VM that is part of VLAN 42, then the switch port that is used by that machine needs to be assigned to that specific VLAN 42. A physical machine can host multiple VMs

which can be on different VLANs, therefore a switch port might be assigned to multiple VLANs (which creates a `VLAN trunk` between the host and the switch port). Whenever a host deploys a new VM or removes a VM, the switch port might need to be reconfigured. Ideally, this can be done in a dynamic and automated fashion, e.g., through network management protocols. Physical switches only pass packets between machines within the same VLAN, which provides an additional isolation mechanism to our VLAN-capable vSwitch that is deployed on all of the hosts.

Address Mapping VLAN tagging does not require any extra address mapping mechanism. VMs discover address information of other VMs using standard discovery protocols as in a non-virtualized environment. However, the vSwitch module that runs on each physical machine learns Ethernet addresses attached to the vSwitch ports by inspecting packets (in the same way as physical switches do) and builds up lookup tables (one table per vSwitch / VLAN) that store information about the location of VMs based on their Ethernet addresses. The vSwitch uses this table to decide if a packet has to be passed to a local VM or onto the physical network to be delivered to a remote machine.

There is no explicit mapping of broadcast / multicast addresses as in the case of EtherIP encapsulation. Instead, physical switches that manage the underlying network infrastructure ensure that broadcast and multicast traffic never crosses VLAN boundaries. Broadcast and multicast packets that are tagged with a VLAN ID are passed to all switch ports that are associated with that particular VLAN, but no other ports. When those packets enter the physical machine that runs our vSwitch module, the packets are only passed into VMs that are attached to vSwitch ports that are assigned to the VLAN matching the ID in the packets.

Requirements Revisited VLAN tagging can be used over arbitrary Layer 3 networks. However, unlike EtherIP a solution based on pure VLAN tagging is limited to a LAN environment and cannot be deployed over WAN links. VLAN tagging highly depends on support from the physical network equipment that is managing the underlying infrastructure. For example, physical switches need to support the VLAN tagging standard that we use when tagging our packets in our vSwitch module (IEEE 802.1Q) and need to be configured to handle tagged packets in order to provide appropriate isolation between VLANs.

A VLAN is a logical network segment and by default network traffic such as broadcast messages or ARP communication is limited to a single VLAN. However, it is also possible to allow communication between (virtual) machines of different VLANs through `Inter-VLAN routing`. There are multiple well-known and standardized solutions to allow this. For example, most physical switches facilitate fast Layer 3 routing between multiple VLANs. This solution offers high performance, but requires that routing policies can be configured on the physical network devices – ideally in an automated fashion. As an alternative, we can also deploy specific VMs that have multiple network interfaces in multiple VLANs and route packets between those – as in the EtherIP encapsulation approach.

VLAN tagging inherently supports communciation with non-virtualized systems. This is because VLAN tagging is a widely used standard that is deployed within infrastructures where physical machines do not run any (network) virtualization software. There is no need for a virtual gateway as in the EtherIP case. Instead, physical switches can be configured to remove VLAN tags from packets when transmitting on

a port where the connected endpoint is not VLAN-capable, and add tags whenever packets are received on that specific port. In that case those endpoints are completely unaware of VLANs, and receive and transmit packets without any VLAN information.

Secure Networking

In addition to EtherIP and VLAN tagging modules, our vSwitches employ network access modules for admission control and optional VPN modules for packet tunneling. Network access control or NAC is an IEEE 802.1X standard for port-based access control [50]. NAC can be implemented by various network devices which in turn force a host (supplicant) go through an authentication process prior to using services provided by the device (e.g., being admitted to the network). All network traffic originating from the supplicant is blocked prior to admission except the NAC traffic itself. NAC requests are received by an access point (authenticator) and forwarded to an authentication server (AS) (e.g., a RADIUS server). The AS runs a choice of authentication protocol specified by the extensible authentication protocol (EAP). It returns the verdict to the authenticator as a result of which access is granted to the supplicant or denied. The NAC module is incorporated into the vSwitch admission process during which the requesting VM is authenticated prior to TVD admission. The choice of which authentication method is dictated by the high-level TVD policies.

Encapsulation and tagging alone do not provide any guarantees on the confidentiality and integrity of the packets / frames that are transmitted on the wire. Without a proper VPN layer, these schemes are only suitable for routed and controlled networks in which the underlying physical infrastructure is trusted, e.g., a virtualized data center. In cases no such guarantees can be given (e.g., over a WAN link), we employ an implementation of IPSec [63] to tunnel VLAN communication in a confidential and an integrity-preserving way. To do so, the VPN module employs the Encapsulating Security Payload (ESP) of IPSec that encapsulates IP packets and applies block encryption to provide confidentiality and integrity. This adds an additional layer of packet encapsulation on top of EtherIP. The optional VPN module is incorporated into the vSwitch.

7.4 TVD Infrastructure

In this section, we present the trusted virtual domain (TVD) infrastructure, the composition of its components to form TVDs and to enforce TVD policies, and describe the management of this infrastructure. We first focus on the static behavior of a secure network virtualization framework that is already up and running. Later, in Section 7.5, we focus on the more dynamic aspects of the framework, including establishment and deployment of the secure virtual infrastructure.

7.4.1 Security Objectives and Policies

Security policies for a TVD can be grouped into two categories: intra-TVD policies that enforce security within a TVD and inter-TVD policies that enforce security across TVDs.

Security within a TVD

Within a TVD, all VPEs can freely communicate with each other while observing TVD-specific integrity and confidentiality requirements. Inside a data center, this

trusted network may be achieved without additional protection mechanisms. On potentially insecure networks, secure intra-TVD communication requires an authenticated and encrypted channel (e.g., using IPsec)[2].

Given a set T of trusted virtual domains, one way of formalizing secure communication requirements as a domain-protection function $P : T \rightarrow 2^{\{c,i,s\}}$, which describes the subset of security objectives (confidentiality, integrity protection, and isolation) assigned to a particular TVD. In TVD context, integrity means that a VPE cannot inject "bad" messages and pretend they are from another VPE. Confidentiality refers to the requirement that two honest VPEs (in the same TVD or different TVDs) can communicate with each other without an eavesdropper learning the content of the communication[3]. Isolation means that each TVD can enforce its security policy independently of other TVDs that share the same infrastructure in a data center. In that respect a TVD must be shielded from policy violations in other TVDs.

Admission control and membership management are crucial aspects of TVDs that help meet these security objectives. VPEs that seek TVD membership may be required to prove their eligibility either periodically or on request. For example, prospective members may be required to possess certain credentials such as certificates or to prove that they satisfy certain properties [96] prior to admission to the TVD in a certain role. The conditions may vary for different roles of VPEs. For example, servers and workstations may have different TVD membership requirements. Similarly, the security requirements on internal machines are usually weaker than for machine acting in a gateway role. Overall, a TVD should be able to restrict its membership to machines that satisfy a given set of conditions for each given role. Some VPEs may be part of more than one TVDs, in which case they would have to satisfy the membership requirements of all the TVDs they are part of. Note that to enable multiple TVD membership, the individual TVD membership requirements must be conflict-free.

One way of formalizing the membership requirements for a set T of trusted virtual domains is to define a function $R : T \rightarrow R$ that identifies the set of unique roles R for a given TVD. We then use a function $M : R \rightarrow 2^P$, where $(P, \leq)$ is a lattice of security properties to identify the required security properties for each given role for that TVD. A machine m with a set p_m of security properties may be permitted to join the TVD t in role r iff $r \in R(t)$ and $\forall p \in M(r) : \exists p' \in p_m$ such that $p' \geq p$. In other words, m is permitted to join t iff there is at least one property of m that satisfies each security requirement of r.

Security across TVDs

Inter-TVD security objectives are independently enforced by each of the individual TVDs involved. To facilitate such multilateral enforcement, global security objectives are decomposed into per-TVD security policies. The advantage of such a decentralized enforcement approach is that each TVD is shielded from security failures in other TVDs, hence enforce domain isolation. Security objectives may take different forms; here, we focus on information flow control among the TVDs and multi-TVD memberships.

An information flow control matrix is a simple way of formalizing the system-wide flow control objectives. Table 7.2 shows a sample matrix for three TVDs: TVD_α,

[2] A network is *trusted* with respect to a TVD security objective if it is trusted to enforce the given objective transparently. For example, a server-internal Ethernet can often provide confidentiality without any need for encryption.

[3] Covert communication channels are outside the scope of this paper.

Table 7.2: Example TVD Policy Specification for Three TVDs

from:		to:		
TVD	Role	TVD_α	TVD_β	TVD_γ
TVD_α	gate	$\mathbf{1}^*$	$\mathbf{0}^*$	$P_{\alpha\gamma}$
TVD_α	internal	$\mathbf{1}^*$		
TVD_β	gate	$\mathbf{0}^*$	$\mathbf{1}^*$	$\mathbf{0}$
TVD_β	internal		$\mathbf{1}^*$	
TVD_γ	gate	$P_{\gamma\alpha}$	$P_{\gamma\beta}$	$\mathbf{1}$
TVD_γ	internal			$\mathbf{1}^*$

Table 7.3: Components of the TVD infrastructure.

TVD COMPONENT	FUNCTIONALITY
TVD Master	A global policy and credential repository per TVD that provisions TVD proxies onto each physical host.
TVD Proxy	A local TVD master delegate, policy enforcer, and credential repository per physical host per TVD.
TVD Coordinator	A local TVD factory per physical host that spawns TVD proxies on request by TVD masters.

TVD_β, and TVD_γ. Each matrix line represents a policy elements that specifies for a given role of a given domain, what other domains it may connect to. The $\mathbf{1}$ elements along the matrix diagonal convey the fact that any connection is allowed while a $\mathbf{0}$ disallows information flow.

While these two policies will later be enforced by allowing or disallowing logical connections, gateway machines can also enforce further application-level policies. These policies P refine the all-flow policy $\mathbf{1}$ by allowing some flows while disallowing others. In order to allow flow from a machine A playing Role $TVD_A/Role_A$ to machine B playing $TVD_B/Role_B$, both corresponding policies need to permit the flow. E.g., if $P_{\alpha\beta} = \mathbf{0}$ in Table 7.2, then $P_{\alpha\gamma}$ and $P_{\gamma\beta}$ should not be inadvertedly $\mathbf{1}$. Otherwise, indirect information flow from TVD_α to TVD_β would be unconstrained, which would contradict with $P_{\alpha\beta}$ and result in a policy conflict.

The TVD policy also includes *multi-TVD membership policy*, where a TVD specifies whether a VPE in a certain role can hold another role in the same TVD or other roles of other TVDs. This is formalized by multi-membership function $M' : R \rightarrow 2^R$ that assigns the subset of roles of other domains that can be assumed by a given VPE. At the time of TVD policy specification, it is important to ensure that there is no conflict between the various policy forms, e.g., if the information flow control policy specifies that there should be no information flow between TVD_α and TVD_β, then the multi-TVD membership restrictions cannot allow any VM to simultaneously be a member of TVD_α and TVD_β.

7.4.2 TVD Components

A TVD infrastructure consists of a single TVD master and multiple TVD proxies that act as a delegate for that TVD on each physical host. We also use a local TVD co-

ordinator per physical platform to spawn all TVDs for that platform during TVD provisioning. Table 7.4.1 lists each component and its role in the TVD infrastructure.

TVD master

The *TVD master* plays a central role in the management and auto-deployment of each TVD. We refer to the TVD master as a single logical entity for each TVD, although its implementation may be distributed. The TVD master is trusted by the TVD infrastructure and the VPEs that are members of the TVD. Known techniques based on Trusted Computing can be used to validate the integrity of the TVD master by verifying its software configuration. The TVD master can be hosted on a physical machine or a virtual machine. In the case of a VM implementation, the PEV architecture proposed by Jansen et al. [55] can be used to obtain policy enforcement and compliance proofs for the purpose of assessing the TVD master's trustworthiness.

The TVD policy is defined at the TVD master by the domain administrator (e.g., the administrator of a data center hosting multiple TVDs agrees on a policy with each customer). The TVD master has the following responsibilities:

1. distributing the TVD policy and other TVD credentials (such as VPN key) to the TVD proxies and informing them of any updates,
2. determining the suitability of a platform to host a TVD proxy (described below) and periodically assessing the platform's continued suitability to host VPEs belonging to the TVD.
3. maintaining an up-to-date view of the TVD membership which includes a list of TVD proxies and the VPEs hosted on their respective platforms.

TVD proxy

On every host that may potentially host a VM belonging to the TVD, there is a local delegate of the TVD master, called the *TVD proxy*. The TVD proxy is the local enforcer of the TVD policies on a given physical platform. At the time of its creation, the TVD proxy receives the TVD policy from the TVD master. Upon an update to the TVD policy (by a domain administrator), renewal of TVD credentials, or refresh of TVD VPN keys at the TVD master, the master conveys the update to the TVD proxies.

The TVD proxies on a given platform are independent. Although TVD proxies are trusted, TVD proxies on the same platform should be sufficiently isolated from each other. For example, a TVD proxy should not be able to access private TVD information (such as policies, certificates, and VPN keys) belonging to another TVD proxy. For improved isolation, each TVD proxy on the platform may be hosted in a separate *infrastructure* VM, which is different from a VM hosting regular services, called *production* VM. On a platform with the Trusted Platform Module or TPM [115], isolation can further be improved by TPM virtualization [13], assigning a separate virtual TPM to each infrastructure VM, and using the virtual TPM as the basis for storing private TVD information.

A TVD proxy must only be able to interact with VMs hosted on the platform belonging to the same TVD. As we describe below, that requirement is enforced by a local TVD co-ordinator. The responsibilities of the TVD proxy are:

Configuration of the Local TVD vSwitch The TVD proxy configures the local TVD vSwitch on a given host based on TVD policy. For example, if the TVD policy

specifies that information confidentiality is an objective, then the TVD proxy enables all traffic through the vSwitch to pass through the VPN module and provides the VPN key to the module.

Caching of the TVD Information The TVD proxy maintains private TVD information such as policies, certificates, and VPN keys.

Status Reports to the TVD Master Upon request or periodically, the TVD proxy provides a platform status report to the TVD master. The report includes information such as the number of VMs belonging to the TVD and their unique addressable identifiers and the current vSwitch configuration. The status report also serves as an *"I am alive"* message to the TVD master, and helps the TVD master to keep an updated list of TVD proxies that are connected to it.

Enforcement of Admission Requirements for VMs into the TVD A VM's virtual NIC is attached to a vSwitch only after the TVD proxies of the switch's and the machine's domain approve this attachment. Note that while the TVD proxy of the switch prevents admission of non-compliant machines, the TVD proxy of the machine prevents outflow into untrusted peer domains.

Continuous Enforcement of TVD Policy The TVD proxy is responsible for continuous enforcement of TVD policy despite updates to the policy and changing configuration of the platform and member VMs. Upon receiving an update to the TVD policy from the TVD master, the TVD proxy may re-configure the vSwitch, and re-assess member VMs' membership to reflect the updated policy. Even without any policy update, the TVD proxy may be required by TVD policy to periodically do such re-configuration and re-assessment.

Local Common TVD Coordinator (LCTC)

The Local Common TVD Coordinator or LCTC is present on every platform (hence, the word *local* in the name) on which a TVD element has to be hosted. The LCTC itself does not belong to any single TVD (hence, the word *common* in the name). The LCTC is part of the minimal TCB[4] on every TVD-enabled platform.

The LCTC is the entity that a TVD master or a system administrator contacts to create a new TVD proxy on the platform. For this purpose, the LCTC must be made publicly addressable and knowledgeable about the identities of the entities that may potentially request the creation.

The LCTC has two main responsibilities, namely (1) creation of new TVD proxies on the local platform, (2) restricting access of TVD proxies only to VMs belonging to their respective TVDs. The LCTC maintains a list of VMs currently hosted on the platform, a list of TVD proxies currently hosted on the platform, and a mapping between the VMs and the TVDs they belong to.

7.4.3 Establishment of the TVD Infrastructure

Preparing a Platform to Host a given TVD

The initial step for establishing a TVD is to create the TVD master (step 0 in Figure 7.5) and initialize the master with the TVD requirements (as formalized above) and the

[4] On a Xen-based platform, the minimal TCB consists of the LCTC, Xen Dom0, the Xen hypervisor, and the underlying hardware.

policy. The output of the step is a TVD object that contains the TVD's unique identifier, i.e., the TVD master's URL and public key.

Once the TVD master has been initialized, the TVD is ready for being populated with member entities, such as platforms and VPEs. A VPE becomes admitted to a TVD after the successful completion of a multi-step protocol (steps 1 and 2 in Figure 7.5):

1. The LCTC on a physical platform creates a *TVD proxy* for the given domain and initializes it with the URL and public key of the TVD master.

2. The TVD proxy sets up a secure, authenticated channel with the TVD master using standard techniques. This includes a validation whether the physical platform satisfies the admission requirements of this specific TVD.

3. The TVD proxy indicates the security and functional capabilities of the physical machine. Using the capability model, the TVD master determines which additional mechanisms must be provided at the level of the virtual infrastructure. For example, if a TVD requirements specification includes isolation and the physical infrastructure does not have that capability, then special (VLAN tagging or VPN) modules must be instantiated within the Dom0 of physical machines hosting VMs that are part of the TVD.

4. The TVD master then replies to the TVD proxy with the TVD security policy (such as flow control policies between VMs belonging to different TVDs hosted on the same physical machine) and additional mechanisms that must be provided at the virtualization level. The TVD proxy also obtains the TVD credentials needed for network security.

5. The TVD proxy then instantiates and configures the required TVD-specific modules (e.g., vSwitch, VLAN tagging module, encapsulation module, VPN module, policy engine, etc.) according to the TVD policy. After this step, the physical machine is ready to host a VM belonging to the TVD.

Adding a VPE to a Requested Domain

Once the platform is ready to host VPEs that are members of a given TVD, a VPE can join the TVD by being connected to the corresponding vSwitch as follows:

1. A VPE requests to join a given domain (identified by master URL and public key) playing a given role.

2. The VPE runs a credential validation protocol with the TVD proxy of the domain that it intends to join. The required credentials usually depend on the role to play.

3. The TVD proxy admits the machine to the domain and connects it with the requested resources such as the domain-internal network.

Connecting a VPE to a Requested Domain

Once a VPE has been assigned to a domain, it may request connection to networks of other domains:

1. A VPE requests to connect a given virtual network adapter to the network of a second domain (identified by master URL and public key).

2. The VPE runs a credential validation protocol with the TVD proxy of the domain of the vSwitch.
3. The TVD proxy of the domain is asked whether connection to this target TVD is permitted.
4. The virtual network of the VPE is connected to the vSwitch.

7.5 Auto-deployment of Trusted Virtual Domains

Figure 7.4 shows the steps involved in automatic deployment of secure virtual infrastructures as TVD configurations. Figure 7.5 shows the steps involved in the establishment and management of a single TVD.

First, the virtual infrastructure topology must be decomposed into constituent TVDs, along with associated security requirements and policy model. Second, a *capability model* of the physical infrastructure must be developed. Capability modeling is essentially the step of taking stock of existing mechanisms that can be directly used to satisfy the TVD security requirements. In this paper, we consider the case where both steps are done manually in an offline manner; future extensions will focus on automating them and on dynamically changing the capability models based on actual changes to the capabilities.

7.5.1 Capability Modeling of the Physical Infrastructure

Capability modeling of the physical infrastructure considers both functional and security capabilities. The functional capabilities of a host may be modeled using a function $C : H \rightarrow \{VLAN, Ethernet, IP\}$, to describe whether a host has VLAN, Ethernet, or IP support. Modeling of security capabilities includes two orthogonal aspects: the set of security properties and the assurance that these properties are actually provided. Table 7.4 lists some examples of security properties and Table 7.5 gives examples of the types of evidence that can be used to support security property claims.

7.5.2 Instantiation of the Right Networking Modules

The TVD proxy uses the instructions given to it by the TVD master to determine the right protection mechanisms to instantiate on the local platform for the TVD network traffic, and accordingly configures the local TVD vSwitch.

Suppose that isolation of TVD traffic is a requirement. Then, VLAN tagging alone would suffice provided the TVD spans only the LAN and the physical switches on the LAN are VLAN-enabled (i.e., it must support IEEE 802.1Q and must be appropriately configured); in that case, a VLAN tagging module would be created and connected to the vSwitch. If the TVD spans beyond a LAN, then VLAN tagging must be used in conjunction with EtherIP encapsulation. In this case, the VLAN tagged packet is encapsulated in a new IP packet and tunneled to the other side, where the original VLAN tagged packet is extracted and transmitted on the VLAN. If VLAN-enabled switches are not available, then EtherIP alone would suffice for isolation.

By itself, EtherIP does not provide integrity or confidentiality of the packets. Hence, when those properties are required, EtherIP is suitable only on routed and trusted networks, e.g., EtherIP would be suitable for traffic between two vSwitches

hosted on different physical platforms that are not connected to the same VLAN switch in a datacenter or corporate environment.

If integrity or confidentiality are required properties and the underlying network is not trusted, then IPsec is used in conjunction with EtherIP and VLAN. In that case, the TVD proxy will create the VPN module, initialize it with the VPN key obtained from the master, and connect it to the vSwitch. Since IPsec only operates on IP packets and not Ethernet or VLAN ones, double encapsulation is needed: EtherIP is used to first encapsulate the Ethernet or VLAN packets, followed by IPsec encapsulation and encryption (using the VPN key).

7.5.3 Inter-TVD Management

Separation of flow control and transport. Transport part is based on today's connection between autonomous systems based on BGP. Flow control is based on firewalls. Flow control part figures out whether a packet can get out to the other TVD and what protection mechanism is needed (e.g., encryption). After flow control, border gateway takes care of routing the packet to the border gateway at other end. Contact master for obtaining capability of the other side gateway. Master-master communication for generating shared key for encryption for inter-TVD traffic.

One firewall for each other TVD, or new rules for another new TVD on the same firewall?

Inter-TVD management deals with the *interchange fabric* for communication between TVDs, enforcement of inter-TVD flow control policies, external zones (IP versus Ethernet), approval of admission requests by TVD-external entities (such as a new VM) to join the TVD, and linking such entities with the appropriate TVD master.

As shown in Figure 7.2, inter-TVD communication can be broadly classified into three types: (1) *controlled* connections, represented by policy entries in the matrix, (2) *open* or unrestricted connections, represented by **1** elements in the matrix, and (3) *closed* connections, represented by **0** elements in the matrix.

Controlled connections restrict the flow between TVDs based on specified policies. The policies are enforced at TVD boundaries (at both TVDs) by appropriately configured firewalls. The TVD master may push pre-checked configurations (derived from TVD policies) into the firewalls during the establishment of the TVD topology. If available, a management console at the TVD master may be used to manually set up and/or alter the configurations of the firewalls. A TVD firewall has multiple virtual network interface cards, one card for the internal VLAN that the firewall protects and one additional card for each TVD that the members of the protected TVD want to communicate with.

Open connection between two TVDs means that any two machines in either TVD can communicate freely. In such a case, the firewalls at both TVDs would have virtual network cards for the peer domain and simply serve as bridges between the domains. For example, different zones in a given enterprise may form different TVDs, but may communicate freely. As another example, two TVDs may have different membership requirements, but may have an open connection between their elements. Open connection between two domains may be implemented using an unlimited number of virtual routers. In a physical machine that is hosting two VMs belonging to different TVDs with an open connection, the corresponding vSwitches may be directly connected. Communication between two TVDs, while open, may be subject to some constraints and monitoring. For example, a TVD master may permit the creation of only

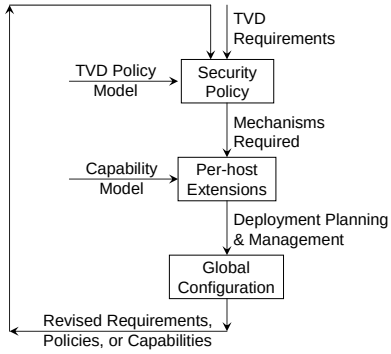


Figure 7.4: Steps in Auto-Deployment of TVDs.

a few virtual routers on certain high-assurance physical machines for information flow between the TVD and another TVD with which the former has an open connection.

A closed connection between two TVDs can be seen as a special case of a controlled connection in which the firewall does not have virtual network card for the peer TVD. In addition to the firewall filtering rules, the absence of the card will prevent any communication with the peer TVD.

Special VMs (e.g., gateways) may have membership in two or more open TVDs simultaneously. Consider a VM that is first a member of TVD α. The VM has one virtual NIC that is connected to the vSwitch of TVD α. Now, suppose that the VM needs to be a member of both TVD α and TVD β simultaneously. For this purpose, the VM needs to have two virtual NICs, one connected to the vSwitch of TVD α and the other connected to that of TVD β. Any VM request to create a new virtual NIC has to be approved by the TVD proxy, which grants the approval only if TVD α is an open TVD and TVD α's policies allow such a dual membership. Initially, the new virtual NIC is connected to the default network, and the VM sends a membership request for TVD β to the local TVD proxy (if present) or a remote TVD master. If TVD β's policies allow for dual membership with TVD α and the VM satisfies other admission requirements for TVD β, then proxy for TVD β will connect the new virtual NIC to the vSwitch of TVD β. At this point, the VM is connected to the VLANs of both TVDs.

7.5.4 Intra-TVD Management

Intra-TVD management is concerned with TVD membership (including mutual authentication), communication within a TVD, and the network fabric (i.e., internal topology) of a TVD. Prior to membership negotiation, mutual authentication requires both the TVD infrastructure (TVD proxy or the master in the absence of a proxy) and the client (i.e., the prospective member VM) to authenticate itself to each other. For TVD authentication, we employ TVD certificates that are issued and distributed for each TVD. For client authentication, we use the IEEE 802.1X standard for network access control (NAC). The latter employs a port-based authentication scheme and a third-party authenticator (e.g., a RADIUS server) to authenticate the VM. In this setting, the TVD proxy acts as the *authenticator* that forwards the request to the *authentication server* and interprets the result. We describe client authentication in detail in Sections 7.5

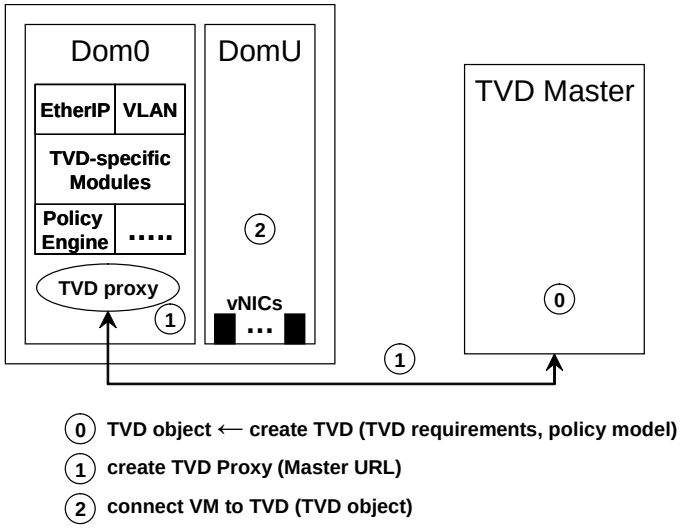


Figure 7.5: Steps in Populating a TVD.

and 7.6.

Intra-TVD policies specify the membership requirements for each TVD, i.e., the conditions under which a VM is allowed to join the TVD. At a physical machine hosting the VM, the requirements are enforced by the machine's TVD proxy in collaboration with networking elements (such as vSwitches) based on the policies given to the TVD proxy by the TVD master. We describe TVD admission control in detail in Section 7.5.

A VLAN can be part of at most one TVD. For completeness, each VLAN that is not explicitly part of some TVD is assumed to be a member of a *dummy* TVD, TVD_{Δ}. Although a VLAN that is part of TVD_{Δ} may employ its own protection mechanisms, the TVD itself does not enforce any flow control policy and has open or unrestricted connections with other TVDs. Thus, in the information flow control matrix representation, the entries for policies, $P_{\Delta\alpha}$ and $P_{\alpha\Delta}$, would all be **1** for any TVD_{α}.

A VM that is connected to a particular VLAN segment automatically inherits the segment's TVD membership. The VM gets connected to the VLAN segment only after the TVD proxy on the VM's physical machine has checked whether the VM satisfies the TVD membership requirements. Once it has become a member, the VM can exchange information freely with all other VMs in the same VLAN segment and TVD (intra-TVD communication is typically open or unrestricted). As mentioned before, a VM can be connected to more than one VLAN (and hence, be a member of more than one TVD) through a separate vNIC for each VLAN.

A VM can become a TVD member either in an active or in a passive fashion. In the *passive membership model*, a VM can be (passively) assigned a TVD membership at the time of its creation by specifying in the VM's start-up configuration files which VLAN(s) the VM should be connected to. Alternatively, in the *active membership model*, a VM can actively request TVD membership at a later stage through the corresponding TVD proxy interface or by directly contacting the TVD master if a TVD proxy is not present on the local platform. When a VM is created, it can use its default

network connection to communicate with the outside world, particularly the local TVD proxy (if present) or a remote TVD master to request TVD membership. In either case, the IP address of the TVD authority must be made available to the VM.

A *closed TVD* is a special TVD that has closed connections with all other TVDs. For example, in a multi-level secure (MLS) military infrastructure, a top secret military domain would be a good candidate for a closed TVD. A closed TVD is shut off from the outside world, and a VM that does not belong to the TVD is part of that outside world. Hence, the active membership model is not applicable for closed TVDs. The TVD master for closed domains maintains a list of pre-approved platforms and will create TVD proxies only on those platforms. Only the system administrator or the TVD administrator can modify the list directly at the TVD master.

Open TVDs are those that are not closed. Both active and passive membership is possible in open TVDs. If the TVD master is directly contacted by the VM, the master checks whether a TVD proxy is already present on the VM's platform. If so, the master instructs the proxy to initiate the admission protocol for the VM. If the TVD proxy is not present, then the TVD master initiates the protocol with the LCTC on the platform to create a TVD proxy.

TVD membership requirements may be checked and enforced on a one-time or on a continual basis. Membership can be a one-time operation in which the requirements are checked once and for all, and thereafter, the VM holds the TVD membership for the duration of its life-cycle. Alternatively, membership requirements can be re-evaluated in an online fashion. The TVD proxy may regularly check whether a VM satisfies the requirements. A session-based scheme may be employed in which a VM is allowed open communication with other TVD members only until the next check (i.e., end of the session).

Policy updates at the TVD master or updates to a platform configuration may result in a platform becoming ineligible to any longer host member VPEs. In that case, the TVD master contacts the LCTC and requests it to destroy the TVD proxy. If the TVD is a closed TVD, prior to the actual destruction of the TVD proxy, members VMs are either migrated to another platform with a TVD proxy or destroyed. If the TVD is a open TVD, the VMs connected to the vSwitch are detached and re-connected to the default network connection. The LCTC sends an acknowledgment to the TVD master after the destruction of the TVD proxy. The TVD master may also rekey the TVD VPN key and distribute the new VPN key to the TVD proxies as a further security measure.

When a platform goes offline (e.g., due to maintenance or network partition), the TVD proxy gets disconnected from the TVD master. In such cases, the disconnected TVD proxy still continues to act as the local TVD authority for the VMs belonging to the TVD. However, the TVD VPEs on the platform are disconnected from the rest of the TVD. The absence of a threshold number of status reports from the TVD proxy causes the TVD master to update its list of TVD proxies. Thereafter, when the platform comes online again, the LCTC on the platform contacts the TVD master indicating that the disconnected platform is online again and contains TVD VPEs that wish to re-connect to the rest of the TVD. That is followed by a prepare phase, similar to the one that happens prior to the creation of the TVD proxy (Section 7.4.2). It is necessary to re-assess the suitability of the platform for still hosting the TVD proxy through the prepare phase, because the TVD policy and the platform state may have changed in the duration when the platform was offline. After the successful completion of the prepare phase, the TVD proxy re-establishes the secure, authenticated communication channel with the TVD master. The TVD proxy obtains the updated TVD policy and other credentials from the TVD master through the channel, and re-configures the networking

Table 7.4: Examples of Security Properties used in Capability Modeling.

Property	Description
TVD Isolation	Flow control policies in place for a TVD.
Network	The actual topology of a virtual network in a physical machine.
Network Policy	Security policies for the network, such as firewall rules and isolation rules stating which subnets can be connected.
Storage Policy	Policies for storage security, such as whether the disks are encrypted and what VMs have permission to mount a particular disk.
Virtual Machines	The life-cycle protection mechanisms of the individual VMs, e.g., pre-conditions for execution of a VM.
Hypervisor	Binary integrity of the hypervisor.
Users	The roles and associated users of a machine, e.g., who can assume the role of administrator of the TVD master.

components according to TVD policy.

7.6 Implementation in Xen

In this section, we describe a Xen-based [11] prototype implementation of our secure virtual networking framework. Figure 7.6 shows the implementation of two TVDs, TVD_α and TVD_β. The policy engine, also shown in the figure, implements the policies corresponding to the TVDs specified in the information flow control matrix of Figure 7.2, i.e., open connection within each TVD and closed connection between TVD_α and TVD_β.

7.6.1 Implementation Details

Our implementation is based on Xen-unstable 3.0.4, a VMM for the IA32 platform, with the VMs running the Linux 2.6.18 operating system. Our networking extensions are implemented as kernel modules in Dom0, which also acts as driver domain for the physical NIC(s) of each physical host. A driver domain is special in the sense that it has access to portions of the host's physical hardware, such as a physical NIC.

The virtual network interface organization of Xen splits a NIC driver into two parts: a front-end driver and a back-end driver. A front-end driver is a special NIC driver that resides within the kernel of the guest OS. It is responsible for allocating a network device within the guest kernel (eth0 in Dom1 and Dom2 of hosts A and B, shown in Figure 7.6). The guest kernel layers its IP stack on top of that device as if it had a real Ethernet device driver to talk to. The back-end portion of the network driver resides within the kernel of a separate driver domain (Dom0 in our implementation) and creates a network device within the driver domain for every front-end device in a guest domain that gets created. Figure 7.6 shows two of these back-end devices, vif1.0 and vif2.0, in each of the two hosts A and B. These back-end devices correspond to the eth0 devices in Dom1 and Dom2, respectively, in each host.

Conceptually, the pair of front-end and back-end devices behaves as follows. Packets sent out by the network stack running on top of the front-end network device in the guest domain appear as packets received by the back-end network device in the driver domain. Similarly, packets sent out by the back-end network-device by the driver do-

Table 7.5: Assurance for Past, Present, and Future States used in Capability Modeling.

Past State	**Description**
Trust	A user believes that an entity has certain security properties.
Mutable Log	The entity provides log-file evidence (e.g., audits) that indicates that the platform provides certain properties.
Immutable Logs	The entity has immutable logging systems (e.g., a TPM-quote [120]) for providing evidence. Since the log cannot modified by the entity itself, the resulting assurance is stronger than when mutable logs are used.
Present State	**Description**
Evaluations	Evaluation of a given state, e.g., Common Criteria evaluations [25].
Introspection	Introspection of a system by executing security tests, e.g., virus scanner.
Future State	**Description**
Policies	By providing policies and evidence of their enforcement, a system can justify claims about its future behavior. e.g., DRM policies and VM life-cycle protection policy.
Audit	By guaranteeing regular audits, organizations can claim that certain policies will be enforced in the future.

main appear to the network stack running within a guest domain as packets received by the front-end network device. In its standard configuration, Xen is configured to simply bridge the driver domain back-end devices onto the real physical NIC. By this mechanism, packets generated by a guest domain find their way onto the physical network and packets on the physical network can be received by the guest domain.

The Xen configuration file is used to specify the particular vSwitch and the particular port in the vSwitch to which a Xen back-end device is attached. We use additional scripts to specify whether a particular vSwitch should use one or both of VLAN tagging and encapsulation mechanisms for isolating separate virtual networks.

The vSwitches for TVD_α and TVD_β are each implemented in a distributed fashion (i.e., spread across hosts A and B) by a kernel module in Dom0, which maintains a table mapping virtual network devices to ports on a particular vSwitch. Essentially, the kernel module implements EtherIP processing for packets coming out of and destined for the VMs. Each virtual switch (and hence VLAN segment) has a number identifier associated with it. The Ethernet packets sent by a VM are captured by the kernel module implementing part of the vSwitch as they are received on the corresponding back-end device in Dom0. The packets are encapsulated using EtherIP with the network identifier field set to match the identifier of the vSwitch that the VM is supposed to be plugged into. The EtherIP packet is given either a multicast or unicast IP address and simply fed into the Dom0 IP stack for routing onto the physical network. The kernel module also receives EtherIP packets destined for the physical host. The module un-encapsulates the Ethernet frames contained in the encapsulated EtherIP packets and transmits the raw frame over the appropriate virtual network interface so that it is received by the intended guest vNIC.

In addition to the kernel module for EtherIP processing, we have also implemented a kernel module for VLAN tagging in Dom0 of each virtualized host. Ethernet packets sent by a VM are grabbed at the same point in the Dom0 network stack as in the case of EtherIP processing. However, instead of wrapping the Ethernet packets in an IP packet, the VLAN tagging module re-transmits the packets unmodified into a

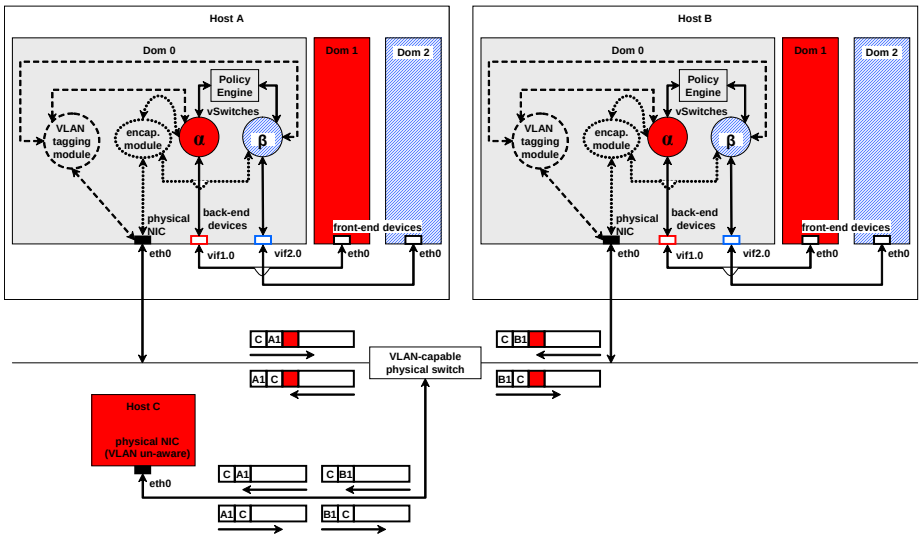


Figure 7.6: Prototype Implementation of TVDs.

pre-configured Linux VLAN device (eth0.α and eth0.β of hosts A and B, shown in Figure 7.6) matching the VLAN that the VM's vNIC is supposed to be connected to. The VLAN device[5] (provided by the standard Linux kernel VLAN support) applies the right VLAN tag to the packet before sending it out onto the physical wire through the physical NIC. The VLAN tagging module also intercepts VLAN packets arriving on the physical wire destined for a VM. The module uses the standard Linux VLAN Ethernet packet handler provided by the 8021q.ko kernel module with a slight modification: the handler removes the VLAN tags and, based on the tag, maps packets to the appropriate vSwitch (α or β) which, in turn, maps them to the corresponding back-end device (vif1.0 or vif2.0) in Dom0. The packets eventually arrive at the corresponding front-end device (eth0 in Dom1 or Dom2) as plain Ethernet packets.

7.6.2 Implementation Issues

Below are some implementation issues we had to tackle in realizing the VLAN and encapsulation approaches.

(1) Some Ethernet cards offer VLAN tag filtering and tag removal/offload capabilities. Such capabilities are useful when running just a single kernel on a physical platform, in which case there is no need to maintain the tags for making propagation decisions. However, for our virtual networking extensions, the hardware device should not strip the tags from packets on reception over the physical wire; instead, the kernel modules we have implemented should decide to which VM the packets should be forwarded. For this purpose, we modified the Linux kernel tg3.ko and forcedeth.ko network drivers so as to disable VLAN offloading.

(2) For efficiency reasons, the Xen front-end and back-end driver implementations avoid computing checksums between them for TCP/IP and UDP/IP packets. We mod-

[5] An alternative approach, which we will implement in the future, is to directly tag the packet and send the tagged packet straight out of the physical NIC without relying on the standard Linux VLAN devices.

ified the Xen code to also handle our EtherIP-encapsulated IP packets in a similar manner.

(3) The EtherIP encapsulation approach relies on mapping a virtual Ethernet broadcast domain to a IP multicast domain. While this works in a LAN environment, we encountered problems when creating VLAN segments that span WAN-separated physical machines. We resolved this issue by building uni-directional multicast tunnels between successive LAN segments.

7.7 Discussion

In this paper, we introduced a secure virtual networking model and a framework for efficient and security-enhanced network virtualization. The key drivers of our framework design were the security and management objectives of virtualized data centers, which are meant to co-host IT infrastructures belonging to multiple departments of an organization or even multiple organizations.

Our framework utilizes a combination of existing networking technologies (such as Ethernet encapsulation, VLAN tagging, VPN, and NAC) and security policy enforcement to concretely realize the abstraction of Trusted Virtual Domains, which can be thought of as security-enhanced variants of virtualized network zones. Policies are specified and enforced at the intra-TVD level (e.g., membership requirements) and inter-TVD level (e.g., information flow control).

Observing that manual configuration of virtual networks is usually error-prone, our design is oriented towards automation. To orchestrate the TVD configuration and deployment process, we introduced management entities called TVD masters. Based on the capability models of the physical infrastructure that are given as input to them, the TVD masters coordinate the set-up and population of TVDs using a well-defined protocol.

We described a Xen-based prototype that implements a subset of our secure network virtualization framework design. The performance of our virtual networking extensions is comparable to the standard Xen (bridge) configuration.

Chapter 8

Public Key Infrastructure

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8.1 Introduction

This chapter outlines a basic design for integration of Trusted Computing (TC) features into a Public-Key Infrastructure (PKI). The adoption of Trusted Computing technologies demands an enhancement of existing infrastructures as well as an adaption of procedures within PKIs. One can identify multiple areas where new development for Trusted Computing is needed:

- First, the design of a trusted platform agent (TPA). Its task is to support initialising, activating and deactivating the TPM security chain under user control. It supports the most important mechanisms and services for creation (or request creation) of keys and credentials related to Trusted Computing. It is capable of communicating with network PKI services.
- Further, a so called "Privacy CA", an entity offering PKI operations (certificate issuance, validation, ...) just like traditional certification authority services, but specialising in Trusted Computing specific tasks. This includes the handling of the TPM Attestation Identity Key credential creation cycle and managing associated request/response messages, keys and credentials. Also, offering services for determination of current status and possible re-evaluation of credentials.
- As a communication protocol between local services (TPA) and network service (privacy CA) the XML Key Management Protocol [129] is employed. It offers functionality to transport traditional PKI operations and enough flexibility for new Trusted Computing specific operations.
- Advanced services are out of scope of this document, only basic services are covered here. The implementation experience of the basic services will lead to a refinement of the services design. Additional services are, e.g., the integration of Subject Key Attestation Evidence (SKAE) extension support, Direct Anonymous Attestation (DAA) as a replacement concept for the privacy CA, and automated policy checking plus validation support. Further, once the XKMS implementation reaches a stable state, an alternative communication protocol will be researched.

8.2 Basic Trusted Computing PKI

A public key infrastructure is a framework enabling authentication, confidentiality and integrity services using public key cryptography. It helps the users of a (public) network to, e.g., authenticate the identity of communication partners and thus establish levels of trust and/or secure communication channels.

The Trusted Computing concept introduces new types of security credentials and procedures. Some fit established structures, some require small adoptions and some represent new concepts.

Associated with the credentials is a life cycle of introducing them to the infrastructure, exchange of information between nodes in the network, (re)validation/evaluation of their information value and finally withdrawal from use.

The new components of a basic Trusting Computing PKI are discussed in the following sections.

8.2.1 EK Certificate

Every Trusted Platform Module (TPM) is (should be) accompanied by a corresponding TPM Endorsement certificate. This certificate contains the public part of the Endorsement Key (EK) pair, which can be viewed as a TPM identity. The private part, called the private Endorsement Key, is stored permanently inside the TPM and can not be retrieved once inserted. The certificate is (typically) signed by the TPM manufacturer and represents an assertion that the specific TPM conforms with the required specifications and the private Endorsement Key is kept safe by a TPM.

Extraction

As per [116] specification a distinct location of non-volatile RAM on the TPM chip is reserved for the TPM EK certificate. Further, the TPM commands to extract non-volatile memory content from the TPM are standardised. Thus, an obvious function of the TPA is to extract the EK certificate. Unfortunately, to this date the only manufacturer to include a TPM EK certificate on chip in every shipped TPM is Infineon.

Creation

If a TPM is shipped without a manufacturer issued certificate, a “late” construction of an EK certificate may be applicable in selected scenarios, e.g., a limited deployment in a department wide setup. Tools for creation of an EK certificate, utilising the real public Endorsement Key of a TPM, are already available from OpenTC partner IAIK. Integration of this functionality into a TPA is aimed for.

Who signs the TPM EK certificate and thus vouches for its integrity is of crucial importance. In a limited deployment scenario a centralised entity can issue homegrown EK certificates as well as offer services for their validation.

In the case of TPM vendor Infineon the necessary certificate chain for validation is freely available for download from the manufacturers homepage. In the case of a self made certificate, the signing authority certificate must be made available and accessible to a validation entity later.

Note that a proof of possession of an EK private key can only be done with a full AIK cycle (see section 8.2.3). This is an intentional limit of the TPM design.

Note also that a TPM EK certificate is the only proof that the corresponding public Endorsement Key actually belongs to a specific type of TPM. Only a certificate signed by a manufacturer (or equivalent important entity) is proof that the referenced TPM is a hardware TPM. Self created certificates may contain an EK public key which actually belongs to a TPM software emulator (e.g., `http://tpm-emulator.berlios.de/`)

Validation

Validation of an TPM EK certificate may be accomplished in multiple steps:

- A local user can read the public EK key from the local TPM and compare it to the one contained in the sample TPM EK certificate. Upon match, one can assume the certificate belongs to the TPM in the local machine.
- If the issuer certificate chain is locally available and the Trusted Platform Agent contains the necessary cryptographic support, a cryptographic validation of the signatures of the certificate chain is possible.
- A thin TPA with minimal footprint may offload certificate verification to a remote service with more resources. In this design the usage of XKMS is suggested (see section 8.4.4).

  Note that the “how” is not as important as the security implications of remote verification. The EK uniquely identifies the TPM, thus, every operation showing the EK to third parties must ensure that the third party can be trusted. Also, security of the communication link with the remote service has to be considered.
- Full validation also requires a check with a PKI of the manufacturer of the specific TPM model (or series), if there are any known conditions affecting the security of the TPM. This infrastructure is out of scope for a basic infrastructure.

Revocation checking is not part of the Basic PKI.

8.2.2 Platform Certificate

The platform manufacturer vouches for the parts of a platform with a Platform Endorsement (PE) Credential. It represents an assertion that the specific platform incorporates a properly certified TPM and the necessary infrastructure according to TCG specifications. There is a requirement for a “root of trust” (CRTM) to be a starting point for building a “chain of trust” and related security measurements are implemented to check the integrity of the platform.

So far no PE certificate is known to have been regularly shipped with a platform. However, a tool to create PE certificates is available from OpenTC partner IAIK. As the

PE certificate is primarily part of the AIK cycle (see section 8.2.3) to be implemented for the basic PKI, the creation of a fake PE certificate with "random" values is aimed for as proof of concept.

8.2.3 Attestation Identity

As the Endorsement Key uniquely identifies a TPM and hence a specific piece of surrounding hardware, the privacy of the user(s) is at risk if the EK is used directly for transactions. As a consequence, the TCG introduced Attestation Identity Keys (AIKs) and associated AIK certificates (standard X509 Public Key Certificates that include extensions defined by TCG), which cannot be backtracked directly to a specific platform. The only entity that possibly knows more details is a trusted third party that issues the AIK certificates, the so called Privacy CA.

AIK certificate creation cycle

In order to create an AIK certificate the following steps are taken:

- The Trusted Platform Agent (TPA, see section 8.3) running on a machine containing a TPM, calls the CollateIdentityRequest function of the Trusted Software Stack (TSS) layer.
- This creates an Attestation Identity RSA key pair and a certification request intended for the Privacy CA.
- The request is transported to the Privacy CA, using proper PKI operational protocols.
- The Privacy CA validates the request content (and included EK and PE certificates). On success it issues an AIK certificate, encrypted with the public EK key of the TPM and thus only readable by the indented recipient.
- The Privacy CA result is communicated back to the TPA.
- The TPA calls the ActivateIdentity function of the Trusted Software Stack, thus unwrapping the AIK certificate.
- The TPA stores the AIK certificate locally.

Summarising, an activated AIK identity comprises a) an "identity" TPM keypair and b) an associated certificate proving that the keypair belongs to a "valid" TPM, vouched for by a Privacy CA entity.

Privacy CA

The role of the Privacy CA (PCA) is of being a trusted third party that works as an anonymiser. For privacy reasons the unique TPM Endorsement Key should only be shown on a "need to know basis". In the concept of the AIK cycle (see previous section) the Privacy CA issues AIK certificates for a "derived" AIK key. This ensures

better anonymity of the EK key holder, but still contains proof of the underlying Trusted Computing supported hardware.

Operation of a Privacy CA is guided by a published policy. It should clearly describe how the relationship EK certificate versus issued AIK certificates is managed. The implementation options for a Privacy CA cover a spectrum from "remember everything" to "know enough for the specific operation, forget everything after completion of operation". Thus, the usage of a specific PCA may be usage scenario dependent.

Implementation of a Privacy CA covers functionality for

- A network front end for receiving/sending requests/responses. The design in this document uses the XML Key Management Standard (XKMS).
- A unit implementing the AIK cycle.
- Local storage. The PCA handles multiple types of certificates. It receives Trusted Computing specific certificates (EK, etc.), it issues AIK certificates and needs foreign certificates for validation (e.g., EK manufacturer certificate chain). The storage must accommodate multiple types.
- A validation unit, capable of determining the status of certificates.

  In the easiest scenario the validation concerns self issued certificates, thus transforming a validation operation to a simple signature check or lookup in local storage. Further, the validation unit should be preloaded with manufacturer certificate chains (e.g., those already available from Infineon), if possible, too.

  The more complex case of actively contacting external entities for missing pieces required for validation is out of scope for a basic PKI.

8.3 Trusted Platform Agent

A PKI requires both server side components, such as certificate authorities, as well as client side applications that provide access to PKI services. In the context of Trusted Computing such a client application is referred to as the *Trusted Platform Agent* (TPA). For wide user acceptance it is crucial that the TPA makes all Trusted Computing related functionality available in a consistent and user-friendly way. Ideally, the TPA is designed and implemented in a modular way that provides an easy integration of additional advanced services later on. Furthermore, in terms of user friendliness the TPA is expected to provide an abstraction of the underlying system concepts that is understandable and manageable for an average user: for this purpose a simple API is provided as well as console commands running on top of it. The TPA largely relies on the services provided by the TSS stack. The overall architecture design of the TPA and the individual system layers is presented in Figure 8.1. Dark grey boxes represent components that will be possibly developed for the Advanced PKI.

The initial basic core functionalities provided by TPA fall in the following categories:

- TPM and platform management. This category includes operations such as TakeOwnership, enabling and disabling the TPM and reading TPM status information.

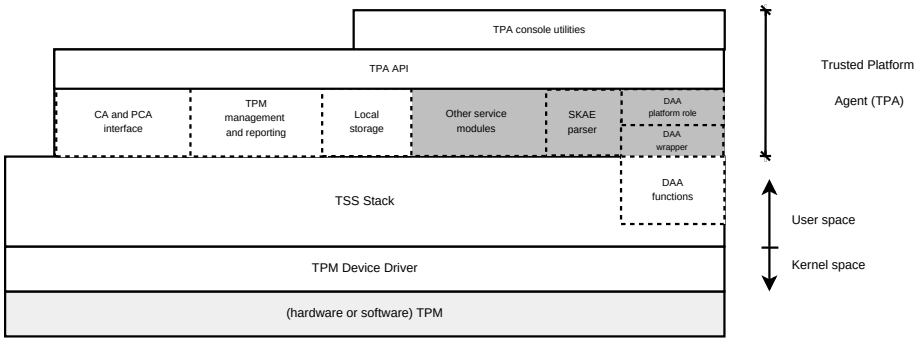

Figure 8.1: Trusted Platform Agent (TPA) and underlying layers

- TC credentials management. This category includes operations needed to manage the life cycle of TC credentials (EK, PE, and AIK certificates) by interacting with TC-enabled authorities. The TPA and the latter communicate through network protocols, XKMS will be used for the first prototype (required extensions will be developed as needed)
  - EK certificate: extraction, creation, validation
  - Platform certificate: creation, validation
  - AIK certificate: creation, validation, reissue, revocation
- Light support for standard X.509 credentials. A simplified support to request a standard X.509 certificate is provided: it is possible to manage certificate with standard profiles using the TC-enabled PKI. This support does not include the interaction with standard PKI authorities; however the interoperability of the issued certificates with existing standard PKIs is guaranteed.
- Local storage for TC-related and standard keys and certificates
- Integrity measurement and reporting. This category includes the following TPM operations: extending PCRs, reading PCRs, activating identities (i.e., AIK certificates) and TPM quote operation.
- API to access all functionalities provided by TPA.

In addition, the TPA can also act as an integration point for a number of other services in the context of Trusted Computing. The main benefit of this approach for the user is that all Trusted Computing related tasks can be done from a single point, the TPA. Adding additional services is facilitated by the modular nature of the TPA. These additional services might include (but are not limited to):

- Management of the DAA communications among the different roles (Trusted Platform, Issuer and Verifier)
  - Standard formats for the exchanged DAA data and messages for using DAA as a standalone protocol or integrated within other protocols
  - A network protocol for using the DAA as a standalone application protocol

- Support for another TC-PKI operational protocol like *Certificate Management Messages over CMS* (CMC) [51] in addition to XKMS.
- Support for the Subject Key Attestation Evidence (SKAE) extension for X.509 credentials.
- A front end for key backup and key migration.
- A user and policy management framework.

8.4 XKMS mapping

A public key infrastructure integrates multiple actors – clients, certification authorities and specialised services. Over the years multiple protocols were developed in the area of PKI and credential management. For Trusted Computing it is necessary to carry traditional PKI services as well as TC specific attributes, queries and data blobs.

XML Key Management Services (XKMS) [129]) is chosen for a first basic Trusted Computing enabled PKI setup, which is in line with the considerations of the TCG in [111] (chap. 6.5.2/p.43) and their recommendation:

"XKMS provides a way to express certificate management function in XML, while providing a wrapper over legacy CA services designed for X.509 certificates. As such, XKMS provides the most attractive solution for credential management for existing CAs in the PKI industry."

XKMS supports four standard registration service functions: Register, Recover, Reissue and Revoke. These offer a wide range of parameters and thus cover the whole life cycle support of credentials.

Further, two key information service functions, Locate and Validate, provide search and status query functionality about credentials deployed in the PKI.

Considering the PKI components outlined in section 8.2 and 8.3, in the following sections a mapping of PKI operations to XKMS specific requests and responses is established and interaction with Trusted Computing usage discussed.

8.4.1 Message Structure

XKMS is an XML based protocol for common PKI operations. The revised edition 2.0 of XKMS [129] reached recommendation status in June 2005. In order to reduce duplicate descriptions in the following sections, the common XML structures of a typical XKMS request and response message are discussed.

Request

The following block outlines the structure of a typical XKMS request:

```
<?xml version="1.0" encoding="UTF–8"?>
<...Request xmlns="http://www.w3.org/2002/03/xkms#"
  xmlns:ds="http://www.w3.org/2000/09/xmldsig#"
  xmlns:xenc="http://www.w3.org/2001/04/xmlenc#"
```

```
  Id="..."
  Service="http://opentc.iaik.tugraz.at/xkms/..."
  ...payload...
  <Authentication>
    ...
  </Authentication>
</...Request>
```

The XML tag name of an XKMS request message always ends in `Request`. Example tag names are `LocateRequest`, `ValidateRequest`, etc. The XKMS XML schema includes the schemata of the XML digital signature standard [130] as well as the XML encryption standard [131]. A good solution is to assign the default XML namespace to XKMS and assign easy recognisable prefixes for the inclusions, as shown above.

Every XKMS message must carry a unique `Id` identifier generated by the originator of the message. Typically this is a random string of at minimum 32 characters (to provide sufficient entropy against attacks).

The `Service` attribute contains the URI of the network service endpoint. For a basic PKI infrastructure the HTTP protocol is sufficient as transport medium. Thus, a XKMS request is mapped to a HTTP POST operation:

```
POST /xkms/... HTTP/1.0
Content-Type: text/xml
Host: opentc.iaik.tugraz.at
Connection: Close
Cache-Control: no-cache
Content-Length: ...

<?xml version="1.0" encoding="UTF-8"?>
<...Request ..... >
```

The path component "`/xkms/...`" is used to distinguish categories of

requests. An obvious mapping would be, e.g., "`.../aik`" for all AIK specific requests and "`.../ek`" for EK related operations. Implementation experience is expected to define useful groupings.

An optional `Authentication` component is employed for operations which are restricted to specific clients or need proof of knowledge of a shared secret. The XKMS standard contains a description of an algorithm to derive a cryptographic key from a secret string (e.g., password). One can then use this key to generate a XML digital signature inside the `Authentication` message component which references the `KeyBinding` type payload of the request. If the validation of the `Authentication` element fails at server side, the response message contains "`ResultMajor=Sender`" with "`ResultMinor=NoAuthentication`".

Response

The following block outlines the structure of a typical XKMS response:

```
<?xml version="1.0" encoding="UTF-8"?>
```

```
<...Result xmlns="http://www.w3.org/2002/03/xkms#"
  xmlns:ds="http://www.w3.org/2000/09/xmldsig#"
  xmlns:xenc="http://www.w3.org/2001/04/xmlenc#"
  Id="..."
  RequestId="..."
  ResultMajor="http://www.w3.org/2002/03/xkms#Success"
  ResultMinor="http://www.w3.org/2002/03/xkms#..."
  Service="http://opentc.iaik.tugraz.at/xkms/...">
  <Signature xmlns="http://www.w3.org/2000/09/xmldsig#">
    ...global message signature of XKMS responder...
  </Signature>
  ...payload...
</...Result>
```

The XML tag name of an XKMS response message always ends in `Result`. Example tag names are `LocateResult`, `ValidateResult`, etc. Note that there also exists a basic `Result` response message. This one is emitted by the server when he cannot properly parse an invalid request and thus cannot determine the more specific type of a request.

In comparison to the XKMS request message the result message contains additional components:

- `RequestId` is a copy of the `Id` of the corresponding request message. It enables a client with multiple XKMS messages in transit to match request-response pairs.
- `ResultMajor` specifies the overall outcome of the request. In case of processing of the request without failure a `Success` result is expected. In the case of an error `ResultMajor` contains an indication who is assumed to be the cause of the error, `Sender` or `Receiver`.
- An optional `ResultMinor` specifies additional details of the result status of a request, if the value in `ResultMajor` can not alone represent all interesting information.

A response by an XKMS service is expected to be always signed. This XML digital signature encloses the whole XKMS message. In order for the client to verify the signature, the public key of the XKMS service must be known on the client side. Typically the public key is shipped to the client in form of a X509 type certificate.

The result received from an XKMS request submitted using HTTP POST typically looks like:

```
HTTP/1.1 200 OK
Date: .....
Content-Type: text/xml; charset=UTF-8
Content-Length: ...
Connection: close

<?xml version="1.0" encoding="UTF-8"?>
<...Result .....
```

8.4.2 RegisterRequest

A XKMS RegisterRequest is used to build a binding of information, typically to a public key(pair). The registration request message contains a prototype of the requested binding.

In the context of Trusted Computing a RegisterRequest may perform the following functions:

Creation of an EK certificate

The TCG infrastructure concept requires the public endorsement key of a TPM accompanied by a certificate. To integrate TPMs (or TPM emulators) without a certificate, a function to create one from a public key is desired.

Structure of a request, including a RSA public key:

```
<RegisterRequest ...>
  <PrototypeKeyBinding Id="....."">
    <KeyInfo ...>
      <KeyValue>
        <RSAKeyValue>
          <Modulus>...</Modulus>
          <Exponent>...</Exponent>
        </RSAKeyValue>
      </KeyValue>
    </KeyInfo>
  </PrototypeKeyBinding>
  <Authentication>
    ...Signature referencing PrototypeKeyBinding...
  </Authentication>
</RegisterRequest>
```

Creation of an AIK identity

The exchange between a client system TPM/TSS and a Privacy CA to create an AIK certificate is almost fully standardised in the TCG specifications. Basically, it comprises a transfer of an encrypted binary blob (namely an array of bytes) to the Privacy CA, resulting in 2 binary blobs as an answer. Unfortunately the features of the XKMS protocol do not allow for an obvious mapping. To prevent early modification of XKMS we decide to transfer the blob information in this case in the `OpaqueClientData` tag. As the name suggests the content of this tag should be opaque to the server, however the gain of experience of getting a running prototype faster has priority. In a later implementation of an advanced PKI the use of, e.g., the XKMS MessageExtension feature for a cleaner solution may be considered.

Structure of the request:

```
<RegisterRequest ...>
  <PrototypeKeyBinding Id="....."">
    <KeyInfo ...>
```

```
        <KeyValue>
          <RSAKeyValue>
            <Modulus>...</Modulus>
            <Exponent>...</Exponent>
          </RSAKeyValue>
        </KeyValue>
      </KeyInfo>
    </PrototypeKeyBinding>
    <Authentication>
      ...Signature referencing PrototypeKeyBinding...
    </Authentication>
</RegisterRequest>
```

The `blob` element containing the binary blob as returned by the CollateIdentityRequest function of the TSS.

Structure of the response:

```
<RegisterResult ...>
  <Signature>...</Signature>
  <KeyBinding>
    <KeyInfo ...>
      <X509Data>
        <X509Certificate>...</X509Certificate>
      </X509Data>
    </KeyInfo>
    <Status StatusValue="http://www.w3.org/2002/03/xkms#Valid"/>
  </KeyBinding>
</RegisterResult>
```

With `blob1` containing the `symCaAttestation` and `blob2` the `asymCaContents` answer of the Privacy CA, to be passed to the ActivateIdentity function of the client TSS.

For a discussion of other `Authentication` possibilities, see also section 8.4.5.

8.4.3 LocateRequest

A XKMS LocateRequest provides a discovery function. It resolves the passed query keybinding and matches request information with local and/or remote data. The answer of a Locate service makes no assertions to any validation criteria. However, a result of a Locate service may be forwarded to a validation service, or, if possible, additional trust verification is done locally.

The following services are useful in a Trusted Computing context:

Query for an AIK certificate

AIK certificates do not contain a subject distinguished name of the certificate owner, but only a `label`, chosen freely at AIK certificate creation time by the client/user. To retrieve a specific AIK certificate a locate request for a specific label name is desired.

An obvious mapping to XKMS would be to use the `X509SubjectName` in the `KeyInfo` portion, however, as some XKMS libraries may check this field strictly for X509 name rules compatibility (and the AIK label specification is less restrictive) this is avoided and the `KeyName` field used instead.

Thus, a query for a specific AIK certificate looks like:

```
<LocateRequest ...>
  <RespondWith>http://www.w3.org/2002/03/xkms#X509Cert</RespondWith>
  <QueryKeyBinding>
    <KeyInfo ...>
      <KeyName>labelOfAikCertificate</KeyName>
    </KeyInfo>
  </QueryKeyBinding>
</LocateRequest>
```

An answer is of the form:

```
<LocateResult ...
  ResultMajor="http://www.w3.org/2002/03/xkms#Success" ...>
  <Signature>...</Signature>
  <UnverifiedKeyBinding>
    <KeyInfo ...>
      <X509Data>
        <X509Certificate>...</X509Certificate>
      </X509Data>
    </KeyInfo>
  </UnverifiedKeyBinding>
</LocateResult>
```

Note that depending on the policy of the Privacy CA the AIK label may not be unique and in the `X509Data` component multiple certificates may be returned.

8.4.4 ValidateRequest

The operations of an XKMS ValidateRequest are similar to a LocateRequest (see previous section), however, the returned status of a binding is evaluated from well defined validation criteria. A validation service returns only information which has been validated by the service. Its validation policy is expected to be publicly available.

In order to validate a specific certificate, it is sent to the service:

```
<ValidateRequest ...>
  <RespondWith>
    http://www.w3.org/2002/03/xkms#X509Chain
  </RespondWith>
  <QueryKeyBinding>
    <KeyInfo ...>
      <X509Data>
        <X509Certificate>...</X509Certificate>
      </X509Data>
```

```
    </KeyInfo>
  </QueryKeyBinding>
</ValidateRequest>
```

The expected result upon positive validation is an `X509Chain`, a certificate chain build from the supplied certificate to a trusted root.

```
<ValidateResult ...
  ResultMajor="http://www.w3.org/2002/03/xkms#Success" ...>
  <Signature>...</Signature>
  <KeyBinding>
    <KeyInfo xmlns="http://www.w3.org/2000/09/xmldsig#">
      <X509Data>
        <X509Certificate>...</X509Certificate>
        <X509Certificate>...</X509Certificate>
        <X509Certificate>...</X509Certificate>
      </X509Data>
    </KeyInfo>
    <Status StatusValue="http://www.w3.org/2002/03/xkms#Valid">
      <ValidReason>
          http://www.w3.org/2002/03/xkms#IssuerTrust
      </ValidReason>
      <ValidReason>
          http://www.w3.org/2002/03/xkms#Signature
      </ValidReason>
      <ValidReason>
          http://www.w3.org/2002/03/xkms#ValidityInterval
       </ValidReason>
    </Status>
  </KeyBinding>
</ValidateResult>
```

The corresponding result message contains the certificate chain as an array of certificates and a `Status` component describing more detailed evaluation results.

In Trusted Computing it is of interest to check the status of EK and AIK certificates. For a basic PKI the XKMS validation message exchange is the same for both cases.

Note that a PE certificate is an attribute certificate whereas XKMS is designed for X509 certificates. An attribute certificate may be included somehow in raw form as array of bytes, but the feasibility of this concept still has to be determined.

Note that it is a policy decision of the service whether the service only validates its own issued certificates or also uses external resources. E.g., validation of an EK certificate may be done locally at the server if the certificate chain is known, however proper validation should also include a revocation check with a manufacturer PKI, if available.

8.4.5 RevokeRequest

An XKMS RevokeRequest manifests the desire to invalidate a previously issued binding. The payload consists of what to revoke, a certificate, etc.:

```
<?xml version="1.0" encoding="UTF–8"?>
<RevokeRequest ...>
  <RevokeKeyBinding Id="...">
    <KeyInfo ...>
      <X509Data>
        <X509Certificate>...</X509Certificate>
      </X509Data>
    </KeyInfo>
    <Status
    StatusValue="http://www.w3.org/2002/03/xkms#Indeterminate"/>
  </RevokeKeyBinding>
  <Authentication>
    ... Signature referencing RevokeKeyBinding...
  </Authentication>
</RevokeRequest>
```

It is expected that this function is always restricted to a specific client population, thus always requires an `Authentication` element.

The response consists of a simple `Success` (or not):

```
<?xml version="1.0" encoding="UTF–8"?>
<RevokeResult ...
  ResultMajor="http://www.w3.org/2002/03/xkms#Success">
  <Signature ...>
    ...
  </Signature>
</RevokeResult>
```

The XKMS options of `Authentication` and/or `RevocationCode` require re-examination under Trusted Computing. Both represent an assurance to the service that one is a valid entity, allowed to withdraw/revoke information from the PKI.

In the case of use of a `RevocationCode` during the RegisterRequest (see section 8.4.2) a code is specified and only if a RevokeRequest supplies the identical code again the revocation is accepted.

The `Authentication` signature can be generated from a shared secret – a password. Usage of a (TPM) private key itself to generate an `Authentication` XKMS signature (effectively a proof of possession signature) is not always feasible in a trusted computing context. The private endorsement key is not available for generic cryptographic operations and the private key corresponding to an AIK certificate is also not designed to be used for arbitrary signing operations.

8.4.6 ReissueRequest

XKMS ReissueRequests are similar to RegisterRequests (see section 8.4.2), the goal being to issue the same item again. The obvious application is to forward an expired certificate and obtain a fresh one of same content, but with a new validity period (the old one getting revoked).

Issues of `Authentication` are similar to those described in section 8.4.5.

Reissuing Trusted Computing related credentials is out of scope for a basic PKI. This point may be revisited when more experiences with certificate life expectancy, usage scenarios and validity periods are available.

8.4.7 RecoverRequest

The XKMS RecoverRequest serves to recover a private key associated with a previously binding. This is only possible if the private key was previously escrowed at the server or server generated. In the context of a basic Trusted Computing infrastructure there is no application for this type of request, as this would invalidate the concept of TPM bound keys, thus can be ignored.

8.5 Open Issues

Design and implementation of a basic trusted PKI for OpenTC highlights multiple issues to be considered. Among them are

- Certificates and issuing authorities require clear and distinct policies. This includes human readable text as well as associated Object Identifiers (OIDs) for automated processing. Only standardisation of these ensures interoperability and spreading of a Trusted Computing PKI.
- The basic PKI outlined in this document assumes XKMS as transport protocol and no specific schema extensions for Trusted Computing. However, even a basic scenario suggests that new URI string definitions for KeyUsage, UseKeyWith etc. would be useful to clearly distinguish TC specific operations from common PKI operations.
- The public documents of the TCG currently only discuss security credentials in X509 certificate format. Some documents however hint at the possibility of future XML based credentials. The inclusion of XML credentials directly in XKMS is a tempting outlook, however the resulting schema extensions and effects on alternative protocols and designs have to be carefully considered.
- At time of this writing the only TPM manufacturer shipping EK certificates with its TPM chips is Infineon. There are no known public platform certificates. There is no known public AIK cycle test. A first basic PKI implementation is hopefully a stimulus for accelerated development, but this highlights that this area is still under major development. Future design adoptions are to be expected.
- The software platform designated to implement this first design on is Linux with its Trusted Software Stack (TSS) called TrouSerS (`http://trousers.sf.net`). At the time of writing this document this is the only freely available fully implemented TSS for the Linux platform. Therefore all experiments and prototyping is using the TrouSerS specific implementation of the AIK cycle. Being heavily tied to low level C structures, level of compatibility of the TrouSerS implementation with other TSS implementations is unknown.

- We have developed prototype implementations of key TCG PKI components. We solved the cryptographic challenges of interacting with a TPM. To our knowledge we are the first to actually demonstrate a working public full PrivacyCA cycle, using TCG style certificates and a dedicated client-server network setup.

Chapter 9

An Efficient Implementation of Trusted Channels based on OpenSSL

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9.1 Motivation

Most of the security sensitive applications on the Internet (e.g., online banking, eCommerce and eGovernment) typically deploy secure channels such as TLS [27] or IPSec [63] to provide secure access to and communication with the corresponding services. These security protocols protect data during transmission and allow to authenticate the endpoints. However, they do not provide any protection from (maliciously) modified software running on an endpoint. More precisely, setting-up a secure channel is currently not linked to the integrity of an endpoint. However, most attacks concern compromising the endpoints by injecting malicious code rather than compromising the secure channel.

This leads to the central problem of today's secure channel protocols: using a secure channel to communicate with an unknown peer opens doors for a wide range of attacks.

Considering a corporate computing at home scenario the following could happen: an employee wants to access from home a corporate's document management server to work on a confidential document. For this purpose he sets-up a secure channel to the company's network and downloads the document. The problem consists in the fact that the employee opened an email attachment containing the executable of a Trojan the day before, which installed itself on the system at this very moment. Now, the attacker that sent the Trojan can access the employee's computer and the document he just downloaded is compromised at the moment it was transferred to the employee's system.

Hence, for the secure provision of digital services over the Internet endpoint integrity is vital. To avert such attack scenarios, information on the communication endpoints integrity or configuration has to be provided in a secure and reliable manner, to enable the peers to judge each other's "trustworthiness" based on the information received.

Reporting integrity information of a remote platform is one of the main goals of Trusted Computing (TC) as proposed by the Trusted Computing Group (TCG, [114]). The basic idea is to securely capture configuration information of the core components of the platform (firmware and software). This information is stored in a cost-effective, tamper-resistant Trusted Platform Module (TPM). The TPM in turn is mounted on the main board of the computing platform and acts as trust anchor. It can sign gathered configuration information and report it to a requesting party. This process is called $attestation$ by the TCG. Additionally data can be stored bound to a specific platform configuration. The TCG calls this mechanism $binding/sealing$ data.

In this paper we focus on the combination of TCG TC functionalities and the TLS protocol to form a Trusted Channel. However, our solution can also be applied with IPSec [63]. Currently, we are considering an implementation of Trusted Channels for this protocol as well.

The central feature of the Trusted Channel is the capability to provide reliable evidence concerning the trustworthiness of a communication partner. Furthermore, by means of a specific system architecture we are able to enforce the security of data not only during transmission but also on the involved endpoints. It has to be pointed out that the linkage of configuration information to the TLS channel is crucial to prevent *relay attacks* where the configuration of a third platform, deemed trustworthy, is relayed by an attacker, acting as Man-in-the-Middle (MitM).

Linking endpoint configuration information to secure channels has been already investigated in the literature [43, 108, 98, 79, 56, 22, 84], often also combined with the TLS protocol because it is the most common protocol used in practice. The TCG also works on this issue in a specific working group [111, 123]. However, none of the solutions so far addresses the problem fully. Some of the approaches only provide an insecure linkage between the secure channel and integrity information, thus MitM attacks seem still possible. Others in turn, have problems concerning their performance in a server environment or required costly acquisitions of, e.g., specific cryptographic hardware (see Related Work in [10]).

In a recent approach [10] a protocol and a generic system architecture for establishing and maintaining Trusted Channels, using TC functionalities and the TLS protocol, was proposed that overcomes most of the shortcomings identified in the afore referenced work. However, the solution in [10] has some deficiencies that our solution aims to tackle: first, some features do not conform to the TLS specification [27], e.g., sending attestation data within the key exchange messages, or including integrity data in session key computation. Changing central message formats or computations of the TLS protocol would result in a time-consuming and costly re-specification process as well as an extensive evaluation of security implications and backward compatibility. Second, [10] supports only RSA key transport, however, *Diffie-Hellman* (DH) can provide perfect forward secrecy of session keys and is also used by a multitude of servers. Third, fundamental functional requirements, like e.g., backward compatibility – to allow communication with systems that do not support integrity reporting – or costs of certification processes are not considered. Certification by, e.g., VeriSign is costly. This means re-certification should take place very seldom, whereas in [10] re-certification would be necessary every time the system is updated, which is the case in practice since systems are updated regularly to overcome security problems or to incorporate new functionalities.

Main Contribution: To overcome the described shortcomings we present a new ap-

proach that bases on [10] but, strictly conforms to the guidelines of the TLS specification and respects central functional requirements listed in detail in Section 9.2.2. Additionally, we focus on a proof of concept implementation of the new handshake protocol to enable the deployment of our approach.

Thus, our main contribution is that our concept (1) fully adheres to the TLS specification and uses existing message extension formats to convey configuration information. To further facilitate a widespread deployment we (2) designed our concept to incorporate functional requirements like, e.g., the possibility to update systems without the need for re-certification, backward compatibility, high-performance system design as well as incurring no additional costs for the users by requiring the use of expensive cryptographic hardware or extensive software adaptations. Apart from that (3) support for all relevant kinds of key exchange methods is provided. Furthermore, we (4) provide stronger forward secrecy of session keys regarding the RSA key exchange method, because keys are held protected by hardware, rendering their disclosure very difficult. Finally, we (5) present a proof of concept implementation of our Trusted Channel protocol.

Outline: In Section 9.2 and Section 9.3, we specify properties and basic terms related to Trusted Channels, followed by our adapted TLS handshake protocol in Section 9.4. Subsequently, we provide a detailed description of the modifications to the handshake messages in Section 9.5. In the Sections 9.6 and 9.7, we first describe the logical architecture and then the implementation of our approach. Finally, in Sections 9.8 and 9.9, we evaluate the whole concept with regard to the security-related as well as the functional requirements enlisted in Section 9.2, concluded by a short Summary.

9.2 Requirement Analysis

In this Section we define the properties of our Trusted Channel concept and derive the requirements necessary to provide those properties.

Adversary Model: The attacker may be a malicious third party, a user or even the administrator of a platform, either eavesdropping the communication between two platforms or controlling one of the peers directly involved in the communication. The adversary is capable to manipulate the software running on a platform, further he can eavesdrop, replace, replay, relay or manipulate data transferred. But, we do not consider sophisticated invasive or non-invasive hardware attacks on involved platforms.

9.2.1 Security Requirements

We adopt the security requirements presented in [10] for a Trusted Channel:

(SR1) Secure channel properties: Integrity and confidentiality of data, freshness to prevent replay attacks, and authenticity both during transmission as well as within the endpoints have to be provided.

(SR2) Authentic linkage of configuration/integrity information and secure channel: Authentic configuration/integrity information must be bound to the trusted channel (i.e., during the establishment and while the Trusted Channel is in place, e.g., the system state changes) to prevent relay attacks.

(SR3) Privacy: Creation and maintenance of the channel should adhere to the least information paradigm, i.e., disclosure of a platform's configuration/integrity information not beyond what is necessary for proper integrity validation. Furthermore, platform configuration information has to be protected against disclosure to a third party.

9.2.2 Functional Requirements

Looking at the wide area of application of TLS on, e.g., servers, desktop-PCs, laptops and infrastructure devices like gateways, all with different functional needs concerning the setup of Trusted Channels, our solution has to adhere to some core functional requirements:

(FR1) Fast deployment support: The alterations to existing software and hardware environments should be minimal and additional concepts introduced should make use of and have to adhere to existing specifications. In addition, all relevant key exchange techniques have to be supported.

(FR2) Minimal costs: The whole approach must not incur additional costs for users like, e.g., for expensive hardware, software or certification.

(FR3) Minimal overhead during handshake: The overhead induced to the handshake by setting-up a Trusted Channel has to be minimal compared to setting-up a common secure channel.

(FR4) Flexible configuration/integrity reporting: It has to be possible to apply different approaches for integrity reporting to support a multitude of differing systems and use-case designs.

(FR5) Backward compatibility: Systems supporting the Trusted Channel approach have to be able to establish conventional secure channels, e.g., to peers that do not provide the means to set-up Trusted Channels.

9.3 Basic Definitions

The underlying system architecture considers client - server ($\mathcal{C}$, $\mathcal{S}$) communication where each involved endpoint may require configuration/integrity information of the other endpoint to be able to judge its trustworthiness.

The evaluation of configuration information is done according to the locally applied *security policy*. If the other endpoint's configuration information conforms to the security policy, this endpoint is considered to be trustworthy. This security policy consists of a set of requirements and guidelines that have to be fulfilled by the platform configuration of the counterpart, e.g., that an appropriate operating system and access control mechanism are in place, etc.

The configuration of a platform is represented by a combination of credentials vouching for security relevant $properties$ of the platform's components (hardware and/or software). Deriving those $properties$ can be done in different ways [96, 47, 37]. The TCG proposes to compute SHA1 [34] hash values over code (software/firmware) for that purpose. The mechanism of deriving these hash values is called $measurement$. These hash values are designated as $digital fingerprints$, since they are used

to unambiguously identify components.

To be able to derive the trustworthiness of a platform we have to compare the digital fingerprints reported by a counterpart to $reference\ values$. In our approach $reference\ values$ represent digital fingerprints provided and signed by a Trusted Third Party (e.g., the distributor or manufacturer of a component). Alternatively, whole certificates can be used to vouch for certain properties of the respective components.

The communication endpoints of our implementation operate based on *compartments*. A compartment consists of one or a group of software components that is logically isolated from other software components. Isolation means that a compartment can only access data of another compartment using specified interfaces provided by a controlling instance.

The set of all security critical software and hardware components of a platform responsible for preserving its trustworthiness is called *Trusted Computing Base* (TCB). Thus, it is crucial to keep the TCB isolated and as small as possible to avoid known problems and vulnerabilities arising along with code complexity.

The central component of the TCB is formed by the TPM, which is currently implemented as a dedicated hardware chip. It offers amongst others a *cryptographic hash function* (SHA-1), a *cryptographic engine* (RSA) for encryption/decryption and signing, a hardware-based *Random Number Generator* (RNG), hardware protected *monotonic counters* as well as some amount of *protected storage*. It provides a set of registers in protected storage called *Platform Configuration Registers* (PCR) that can be used to store hash values. The value of a PCR can only be modified in a predefined way[1]. Protected storage is also used to store certain security sensitive keys, e.g., *Attestation Identity Keys*[2] (AIKs) or the $Storage\ Root\ Key$[3] (SRK).

To improve security of the common TLS protocol, we move all security relevant operations like, e.g., encryption, signing and the handling of credentials to the TCB, whose code is protected against a wide range of attacks (see Section 9.6) running in separate memory space and only accessible via interfaces. The protocol implementation remains in user space, because there is no need to protect it.

9.4 Adapted TLS Handshake

In this Section we describe the high-level adaptations we introduce to the TLS handshake protocol. We focus on the Diffie-Hellman Ephemeral (DHE) key exchange method in a mutual attestation scenario, but our design supports all common TLS key exchange types [27]. Structures we added or altered are depicted in bold text in Figure 9.1. A detailed description of attestation structures is given in Section 9.5.

Negotiating Security Parameters: To set up a Trusted Channel, $\mathcal{C}$ starts the necessary TLS software and sends a $ClientHello$ message to $\mathcal{S}$ that answers with a corresponding $ServerHello$ message. Using those hello messages the two parties involved in the communication negotiate the attributes of the Trusted Channel they want to establish.

[1] $PCR_{i+1} \leftarrow \mathsf{Hash}(PCR_i|x)$, with old register value PCR_i, new register value PCR_{i+1}, and input x (e.g., a SHA-1 hash value). This process is called *extending* a PCR.

[2] These are specific signing keys defined in the TCG specifications that can be used to authenticate a user and/or his system. They are kept securely inside the TPM and can only be used for signing stored measurement values or certifying other non-migratable keys [114].

[3] This key is kept inside the TPM as root for the whole key hierarchy [114].

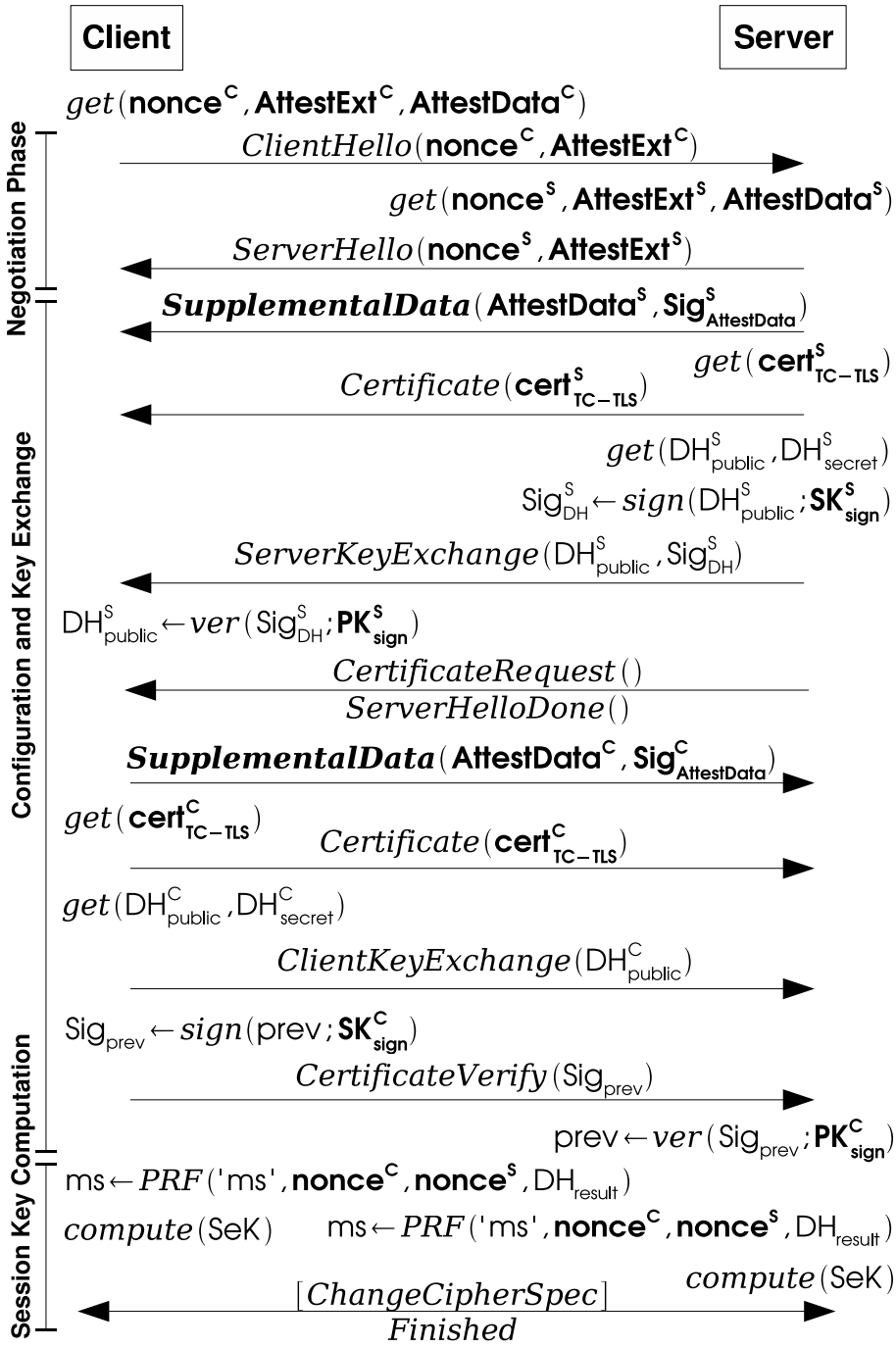

Figure 9.1: Adapted TLS DHE-RSA Handshake

In contrast to the common TLS design the nonces sent in the hello messages are taken from a RNG seeded by the TPM at boot-time[4]. Furthermore, each side includes an *Attestation Extension* ($AExt$), that is used to specify details concerning the configuration/integrity information that will be exchanged.

Configuration and Key Exchange: Each peer provides evidence related to its configuration and integrity. For this purpose we use additional $SupplementalData$ messages as defined in Internet Engineering Task Force (IETF) RFC4680 (cf. Section 9.5, [105]). Thus, $SupplementalData$ messages are composed to transfer *Attestation Data* (aD) representing configuration/integrity information. aD is signed using a secret key (SK_{sign}) for authentication and integrity protection of configuration information. Following the $SupplementalData$ message, each side provides a certificate ($cert_{TCLS}$) including the respective public key (PK_{sign}) used to verify this signature.

Subsequently, $DHvalues$ (DH_{public}, DH_{secret}) are computed on both sides. DH^{S}_{public} is signed using SK^{S}_{sign} to provide authentication evidence. $\mathcal{S}$ then sends DH^{S}_{public} and a signature (Sig^{S}_{DH}) to $\mathcal{C}$ within the $ServerKeyExchange$ message. $\mathcal{C}$ computes its own values and sends DH^{C}_{public} to $\mathcal{S}$ using the $ClientKeyExchange$ message. The following $CertificateVerify$ message is used to prove the possession of SK^{C}_{sign}, and to authenticate DH^{C}_{public}, by signing a digest over all previously exchanged handshake messages ($prev$) using SK^{C}_{sign} [27].

Session Key Computation: Following $CertificateVerify$, the TLS $master\ secret$ (ms) is computed on both sides using $nonce^{C}_{TPM}$, $nonce^{S}_{TPM}$, a string indicating that this is a ms, and the result of the final *Diffie-Hellman* computation as input to a $pseudo\ random\ function$ (PRF). Subsequently, the Session Key (SeK) is derived from the ms on both peers [27, p.24]. At last, the handshake is finalized by the $ChangeCipherSpec$ protocol and final $Finished$ messages. These $Finished$ messages are already encrypted using SeK, thus, a failure in key exchange would be noticed.

9.4.1 State Changes

Since state changes might happen on both peers while the Trusted Channel is in place, we provide the possibility to exchange updated integrity information in a short re-handshake. This re-handshake is triggered when a state change, e.g., the execution of another software in the same compartment, occurs on any side, and if the corresponding state monitoring option was selected in $AExt$ (cf. Section 9.5.2)[5]. In case the parties agreed on state change notification during the initial handshake, the following procedure takes place: If a state change happens on one platform access to the SeK can be blocked and/or access to data belonging to the session is restricted depending on the security policy of the application. Both sides are notified using $HelloRequest$, $ClientHello$ and $ServerHello$, respectively. The updated integrity information for validating the new configuration is securely transmitted to the counterpart encrypted using SeK and included in a data structure called *State Change Extension* ($SCExt$) (cf. Section 9.5.2). Subsequently, a TLS resume message flow takes place [27]. After the short

[4]We use the TPM as source for random values, because its RNG is considered as true random generator in contrast to the pseudo-random generator implemented in OpenSSL. The advantage of this feature becomes obvious with respect to the recently discovered security weakness in the OpenSSL RNG in Debian Linux systems [112].

[5]Changes inside the compartments can be detected using the Integrity Measurement Architecture (IMA, [103]) proposed and implemented by IBM.

TLS resume handshake the new session key SeK' is computed and the communication can continue, or the channel is torn-down because the requirements of the security policy of the peer are not fulfilled any longer.

9.5 Detailed Description of Attestation Data Structures

In this section we introduce credentials and extensions to TLS handshake messages we use to set up a Trusted Channel, followed by an example of their usage.

9.5.1 Key Exchange Types and Certificate Elements

In a common TLS handshake certificates are used to authenticate the peers. There are different certificate types used for different key exchange methods. Depending on the type of certificate chosen or key exchange method supported, we add extensions to the certificates. The certificate extensions and credentials we define are necessary to bind the TLS channel to the endpoints whose configuration is reported, and to be capable of proving that a certain TCB is in place. These new credentials are held in an environment protected by TC mechanisms. Therefore, we act on the assumption that these additional security measures justify stronger security assumptions concerning the storage and usage of those credentials. In the following paragraphs we explain their creation, storage and interdependency.

SKAE Key (K_{SKAE}) and SKAE: The $non\text{-}migratable$[6] asymmetric key pair K_{SKAE} (PK_{SKAE}, SK_{SKAE}) is created after an AIK (K_{AIK}) has been certified and installed. Its private part SK_{SKAE} is sealed to a specific TCB using the SRK ($K_{storage}$) and never leaves the TPM unencrypted. We make use of the *Subject Key Attestation Evidence* ($SKAE$) as proposed by the TCG [110]. In contrast to the intended purpose, we use the $SKAE$ as standalone element within our handshake, but we also foresee the possibility to include it in a X.509 certificate. The $SKAE$ basically consists of a `TPM_CERTIFY_INFO2` structure representing the TCB configuration that has to be in place during key release (including a digest of PK_{SKAE} [118, p.96]) and a signature over this structure (Sig_{SKAE}) by a K_{AIK}. Additionally, links to reference values can be provided. The $SKAE$ can vouch that K_{SKAE} was created by a *Trusted Platform* that conforms to the TCG specification [114] and that a certain TCB configuration has to be in place during release because of sealing.

Secure Encryption Key (K_{enc}) and Secure Signature Key (K_{sign}): We introduce the asymmetric key pairs *Secure Encryption Key* K_{enc} (PK_{enc}, SK_{enc}) and *Secure Signature Key* K_{sign} (PK_{sign}, SK_{sign}), that are considered long-lived and usable for client compartments that wish to establish a Trusted Channel to a remote party. They are created inside the TCB and sealed using $K_{storage}$.

Depending on the key exchange method supported and the respective certificate, PK_{enc} and/or PK_{sign} are included in the common *TLS certificate* ($cert_{TCLS}$) as encryption or signature key (cf. Figure 9.2). Thus, either the public part of K_{enc} or K_{sign} are put into the public key field of the X.509 certificate. In case of RSA and DH_RSA key

[6] The private part of an asymmetric key pair labelled $non - migratable$ never leaves the TPM unencrypted [114]

Key Exchange / Certificate Element	RSA	DH_RSA	DHE_RSA
Certificate Type	Encryption Certificate	Encryption Certificate	Signature Certificate
Key Type	Encryption Key	DH Public Values	Signature Key
Certificate Extension Type	Signature Key Extension	Signature Key Extension	

Figure 9.2: Handshake Types and X.509 Certificates

exchanges, K_{sign} is included as Signature Key Extension to the TLS certificate. Its X.509 format, including the use of extensions, is specified in [48]. The signature key is needed in these handshakes to provide the binding between integrity information and the endpoints. This is not feasible using the encryption key in those handshake types, and using a single key for encryption and signing is considered insecure.

K_{enc} and/or K_{sign} must be used for client and server authentication during the TLS channel setup to guarantee the binding of the secure channel to the integrity state of the endpoints. Therefore, the usual TLS authentication scheme will be used for server authentication and its $cert^{S}_{TCLS}$ will likely be signed by a CA like e.g. VeriSign. In contrast, the TLS client authentication mechanism, optional for a standard TLS channel, must be used to guarantee the binding through $cert^{C}_{TCLS}$ even if the actual and reliable authentication of $\mathcal{C}$ might not be needed. Therefore, in this case $cert^{C}_{TCLS}$ can be self-signed.

SK_{enc} and SK_{sign} are loaded and decrypted during the start of the platform and kept inside the TCB. Subsequently, we need K_{SKAE} to authenticate K_{enc} [7] and K_{sign} [8]: By signing PK_{enc} and PK_{sign} using SK_{SKAE} we provide twofold evidence: that the TCB identified by the $SKAE$ was in place during the signature[9] and it is a statement from that TCB about K_{enc} and K_{sign} like: "I certify that K_{enc}, K_{sign} are correctly treated, i.e., when decrypted, the keys are kept secret by myself". If the verifier of the $SKAE$ trusts the TCB attested by it, then the verifier can also trust the TCB's statement about the correct treatment of K_{enc} and K_{sign}. Therefore, the `TPM_Sign()` function is applied to sign K_{enc}'s and K_{sign}'s public parts with SK_{SKAE} at boot time[10]. The resulting signature Sig_{BSEK} is held in the TCB memory space.

9.5.2 Extensions used in the TLS handshake

The extensions to the TLS protocol that will be introduced in the following paragraphs are necessary to trigger and negotiate the exchange of configuration information as well as for the transport of the additional configuration/integrity data. Extensions to the TLS handshake protocol can be small data chunks added to the Hello messages [27] or completely new handshake messages. These extensions are explicitly foreseen by the TLS

[7] In case DH_RSA key exchange was chosen, the DH parameters included in the certificate are signed.

[8] It would also be possible to use K_{SKAE} instead of K_{sign}, K_{enc} for signing/encrypting during the handshake. But then the involvement of the TPM every time a Trusted Channel is set up would be necessary. This would result in a significant performance loss, especially in connection with server systems. Using K_{AIK} directly for this purpose is not allowed by the TCG specification.

[9] The TCB must check *digest at creation* and *digest at release* of the stored key data objects [118, p.89] before signing them with SK_{SKAE} to be sure that they were not compromised by a former TCB.

[10] This has to be done at every system boot because otherwise the TCB update mechanism presented in Section 9.6.2 could be compromised. Thus this signature is held in volatile memory.

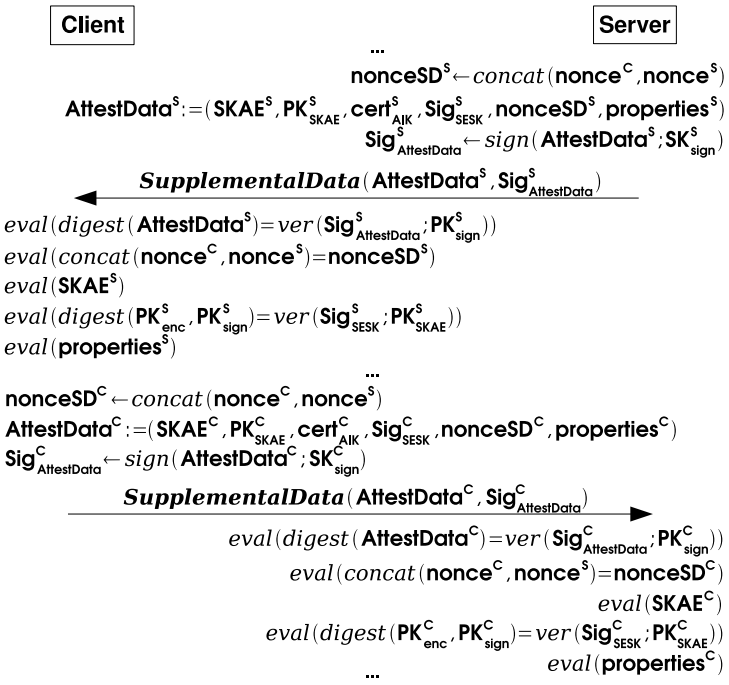

Figure 9.3: Supplemental Data Message Creation and Evaluation

specification to deal with advancements and changes in communication infrastructures. There already exist several extensions to the TLS protocol, e.g., for sending client certificate URLs or explicit server name indication [95]. The basic concept for extending TLS with an additional handshake message is described in RFC4680 available from the Internet Engineering Task Force (IETF) Networking Group [105]. This RFC defines the additional $SupplementalData$ handshake message envisioned to carry additional generic data, whose format must be specified by the application that uses it, and whose delivery must be negotiated via Hello message extensions.

Hello message extensions: $Attestation\ Extension$ or $State\ Change\ Extension$ ($AExt$, $SCExt$) are transmitted within the $ClientHello$ and $ServerHello$ messages. The first one is used in the initial handshake to negotiate which side ($\mathcal{C}$ and/or $\mathcal{S}$) has to attest to its state, the type of attestation and state-monitoring supported or if privacy of configuration information is desired. $SCExt$, in turn, is used to inform the peer of a state change on the counterpart and to transport configuration data in a re-handshake (cf. Section 9.4.1).

Supplemental Data Message Creation and Evaluation: The $SupplementalData$ message includes the $SKAE$, PK_{SKAE}, $cert_{AIK}$, Sig_{BSEK}, a concatenation of the nonces sent in the TLS Hello messages and $properties$ (see Section 9.3), depending on what kind of attestation and key exchange was chosen. Furthermore, a Signature Sig_{aD} on aD is appended. This signature is needed to bind the aD structure to the respective secure channel endpoint. In Figure 9.3 we show how aD is composed, while the $properties$ field is considered as black box, since its values vary depending on the attestation concept chosen.

In our proof of concept prototype the $properties$ data consists of measurements representing the images of the client compartments using the TLS Trusted Channel extracted from a *Configuration Data Structure* (CDS) and corresponding signed $reference\ values$ (see Section 9.3). The CDS itself holds a list of measurements of the binary images of all client compartments running on top of the TCB.

Determination of endpoint trustworthiness: The trustworthiness of the peer's TCB is determined by evaluating Sig_{SKAE} (cf. Section 9.5.1) and the TCB measurement included in $SKAE$ using $reference\ values$ provided by trusted third parties. To verify the validity of K_{sign}, Sig_{BSEK} is checked. Then the linkage between secure channel endpoint and aD is verified by inspecting Sig_{aD}. Freshness of aD is guaranteed comparing $nonceSD$ to the nonces sent the hello messages. Finally, the trustworthiness of compartments running on top of the TCB is determined in a next step by evaluating CDS using either additional $reference\ values$ provided by the peer within the $properties$ data field or by trusted third parties.

Also other concepts of attestation are supported. If, e.g., the TCG attestation mechanism should be used, a digest of $nonceSD$, PK_{enc} and PK_{sign} is given as external data to the `TPM_Quote()` function of the TPM [114]. This is done to provide freshness of TCG attestation data (aD_{TCG}) and to replace Sig_{BSEK}. $SKAE$ is not needed here because the AIK is used to sign the relevant values inside the TPM.

Since TLS handshake messages are usually sent in clear text, in case $privacy$ of attestation information is desired by one of the communication partners, no $SupplementalData$ messages are sent within the first handshake. Subsequently, a second handshake is performed directly after the first one to exchange attestation information encrypted using the session key negotiated in the previous handshake [105].

9.6 Generic System Architecture

Our generic system architecture is based on security frameworks as proposed, e.g., in [97], [102], and consists of an *Application*, *Trusted Service*, *Virtualization* as well as *TC-enabled Hardware Layer*. We kept our approach generic, thus it is possible to implement/integrate the components in different systems also on common operating systems like, e.g., Linux or Windows. But, if these monolithic OSs are applied, some constraints have to be considered when looking at the security of such implementations, because in general they are not capable to ensure strong isolation of processes and corresponding data.

TC-enabled Hardware Layer: The hardware layer has to offer TC extensions that are conforming to the relevant TCG specifications (e.g., [114]). This essentially means that it comprises a TPM chip and a compatible BIOS.

Virtualization Layer: The virtualization layer offers and mediates access to central hardware components like, e.g., CPU and MMU. These tasks can be performed by many kinds of virtualization techniques, namely hypervisors, microkernel approaches or a common OS running a virtualization application, e.g., VMware [125].

Trusted Service Layer: This layer consists of security services [10] and provides interfaces to the Application Layer. It also mediates and monitors access to virtualized hardware resources. Subsequently, we briefly describe the main components of the

TCB in our approach:

- *Trust Manager* (TM) provides functionalities used for establishing Trusted Channels. To be able to provide this functionality TM bundles multiple calls to the TPM into a simple API for calling instances. Thus, it offers functionality to generate keys, bind/unbind, seal/unseal, certify these keys or to report the current measurement values of the TCB stored inside the TPM. The keys used in the initial handshake are computed and held by the TM. They never leave the TCB.
- *Compartment Manager* (CM) is responsible for starting and stopping compartments. It measures the compartment code when starting it and assigns a locally unique ID to this compartment. This ID as well as the measurements are reported to the *Integrity Manager* (see below).
- *Integrity Manager* (IM) stores the compartment's $properties$. In our approach this means appending the measurements reported by CM together with a unique ID to CDS (cf. Section 9.5). IM keeps the CDS secure by storing it inside the TCB's memory space and provides it to other TCB components.
- *Policy Manager* (PM) stores platform and application policies and provides them to other components of the TCB when needed[11].
- *Storage Manager* (SM) handles persistent data storing for the different compartments.

Application Layer: In this layer the applications run in isolated compartments. This can be either applications running directly on top of the underlying TCB or whole OSs.

9.6.1 Trusted Initialization

To be able to attest a platform's configuration, its hard- and software components are measured reliably and those measurements are stored securely. An ongoing measurement process is effected originating from the *Core Root of Trust for Measurement*[12] that initiates the measurement process up to and excluding the Application Layer. Every component that has to be loaded during the boot process is measured before passing control over to it.

Consequently, a *Chain of Trust* (CoT) is established and the TCB is measured reliably. These measurements are stored inside the TPM and represent the *static configuration* in our approach, because it must be only modifiable with a subsequent reboot. After the boot process platform monitoring is conducted by CM. Thus, CM extends the CoT when a client compartment is loaded that runs on top of the TCB. CDS that reflects the platform configuration is maintained by IM. The configuration of compartments that run above the TCB represents the *dynamic configuration* because we allow state changes to happen.

In our approach the CoT is initially built-up until the TCB is loaded and running. To be able to provide support for Trusted Channels a RNG, seeded using `TPM_GetRandom()` at boot-time, provides random. After the system has booted-up and the TCB is in place, TM unseals K_{enc} and K_{sign} and signs their public parts with SK_{SKAE}. TM now holds K_{enc}, K_{sign}, PK_{SKAE} and Sig_{BSEK}. Thus, the system is

[11] This component is not implemented yet, for the prototype we applied fixed policies.

[12] This is a small piece of code initiating the measurement process at the very beginning of the boot process. Usually, this code is located within the BIOS.

initialized and ready to set up a Trusted Channel.

9.6.2 TCB Update Management

We want to be able to change the TCB configuration without requesting a new TLS certificate every time this is done. This is an important issue because on the one hand re-certification of the TLS endpoint and its keys is expensive and on the other hand we want the keys to be securely bound to a specific platform configuration to be able to prove that they have not been compromised.

Using non-migratable keys during the handshake would be the safest way to protect keys from being compromised but resealing such keys to a different TCB is not possible [119]. Therefore, we present a procedure that allows the update of the TCB without compromising the TLS keys and thus preserves the validity of the TLS certificate.

The problem here is that a system in an updated state (and its peer) has to be able to judge its former state, because otherwise K_{enc} and K_{sign} may have been compromised and a new certificate is needed. To achieve this, we keep a secure $changelog$. The $changelog$ holds names and hash values of components that have been installed or removed. Additionally it contains a link to a certificate by a trusted third party (e.g., manufacturer) that vouches for these values.

The update process starts with unsealing SK_{sign} and SK_{enc}. Then the new $package$ [13] is downloaded from a trustworthy entity together with $cert_{update}$ containing the hash value of the component after installation. Then, TM computes the foreseen configuration of the platform after the installation using the hash value comprised in the certificate replacing the values of the removed component in the CDS and it updates $changelog$ accordingly. Subsequently, SK_{sign} and SK_{enc} are sealed to this state. In a last step the new $package$ is installed. After this process the platform has to be rebooted to let the changes take effect.

After the platform is initialized again, a new K_{SKAE} and related $SKAE$ have to be created. From then on the new SK_{SKAE} is used to sign PK_{enc} and PK_{sign} creating Sig_{BSEK}.

9.7 A Trusted Channel Implementation with OpenSSL

As basis for the implementation of the Trusted Channel we chose the Xen Hypervisor [128]. The components – TM, SM, IM, and CM– mentioned in Section 9.6 have been developed within the OpenTC [86, 68] and EMSCB [36] projects. We used these services and extended them where needed.

9.7.1 Implementation Architecture

In order to confine the size of the Trusted Computing Base (TCB), we designed the Trusted Channel split into two groups of components.

One group (the $TLS\ backend$), minimal and part of the TCB, performs the security critical operations: encryption/ decryption, MAC calculation/verification and the

[13] This could be package structures as used by several Linux distributions, e.g., rpm or deb.

management of the TLS session are performed by TM, while the storage of sensitive data like keys and certificates is done by SM [14].

The actual protocol implementation (the $TLS\ frontend$), instead, runs in the same compartment as the application and is not needed to be trusted, thus its size is not important. Applications already using the TLS implementation we chose to enhance (OpenSSL) must be slightly modified to be able to set up a Trusted Channel instead of the standard TLS secure channel.

But, our architecture is also capable to support the provision of Trusted Channels as a service, running in a separate compartment and implemented as proxy. If so, the application directly bound to the adapted TLS protocol is the service providing the Trusted Channel to other applications.

We identified at least two types of interaction between a Trusted Channel service and a generic application: explicit or implicit invocation. In the first case the application explicitly requests the service (i.e. the proxy) for setting up a Trusted Channel. A convenient implementation could use the SOCKS[15] protocol as a carrier of the necessary calls. In case of implicit interaction, the application is unaware of the setup of the Trusted Channel. A communication policy, set within the TCB, enforces transparently the redirection of the application's communications to the Trusted Channel service. A convenient implementation could rely on a transparent proxy, like the one implemented by SQUID [106].

With both explicit and implicit invocation, a scheme using a pair of proxies on both sides can be set up, thus implementing a tunnel via Trusted Channel to carry the end-to-end communication. With the explicit invocation, it is also possible so set up a scheme with a proxy only on the client-side, directly connecting via Trusted Channel to the server.

9.7.2 Enhancements to OpenSSL

OpenSSL [87] is a multi-platform and widespread software toolkit implementing cryptographic operations, SSL [53] and TLS [27] protocols, the encoding/decoding of X.509 [48] certificates and of other PKI-related formats like PKCS [60] standards. It consists of two shared libraries (`libssl` and `libcrypto`) implementing all the features and a command line tool (`openssl`) wrapping them. The libraries can also be directly used by generic applications.

OpenSSL also offers the possibility to delegate the execution of (a subset of) cryptographic operations to a separate module called engine. This is a shared library, with a well-known interface, which can be dynamically loaded at run-time and used by OpenSSL's core libraries: more than one engine can be used at once. An engine can be implemented in software or the library can be just the driver for a hardware cryptographic device.

We built the Trusted Channel around three different enhancements to OpenSSL, represented in the right-hand part of the Figure 9.4 and described in the following.

TLS protocol extensions: The stable version of OpenSSL (0.9.8x) does not implement any extension while the development version (0.9.9x) only includes hard-coded

[14] The trusted initialization described in Section 9.6.1 is realized using the TrustedGRUB (*tGRUB*) boot loader [4]

[15] SOCKS is a well-known protocol for the communication with proxies [73]. The proxy is then the application directly using the adapted TLS and running as service.

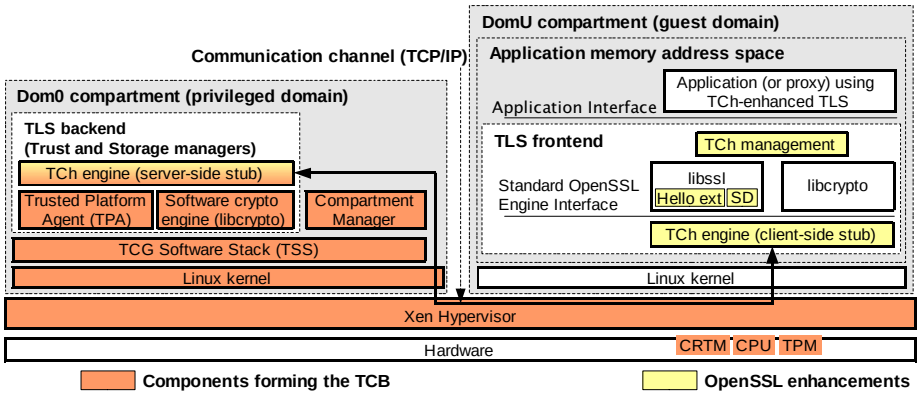

Figure 9.4: OpenSSL enhancements and Proof of Concept prototype

support for the extensions defined in [95] to $ClientHello$ and $ServerHello$ messages. We implemented a mechanism to easily add generic and application-defined extensions to these messages, usable to trigger the delivery of $SupplementalData$ handshake message [105], also newly implemented because not natively supported. These are the only direct enhancements to OpenSSL, implemented as patch to `libssl`'s code, which expose an API for registering new Hello extensions or data for $SupplementalData$ via callback functions.

Trusted Channel management library: This is a completely new module realizing the Hello extensions $AExt$ and $SCExt$ (on top of the enhanced `libssl`) and the logic of the Trusted Channel's specific operations. It also handles the parsing and the validation of the credentials received during the handshake (PK_{AIK}, PK_{sign} and PK_{enc}) and it manages the validation of the attestation data (i.e. aD and Sig_{aD}) carried through the $SupplementalData$ message. Finally it provides the application with an interface to set up the Trusted Channel.

TLS backend via split engine: To implement our concept of delegating the security critical operations, we implemented an OpenSSL engine split into two parts. The shared library implementing the engine interface runs in the application (or proxy) memory space together with the Trusted Channel management library and `libssl`: they all form the $TLS\ frontend$. The latter requests the execution of critical operations over a communication channel at the $TLS\ backend$, which actually provides the functions needed by the TLS protocol and runs in a different compartment as part of the TCB.

9.7.3 Proof of Concept (PoC) Prototype

Figure 9.4 shows a PoC prototype of our OpenSSL-based Trusted Channel implementation built upon Xen.

The $TLS\ backend$ runs in Domain0[16] – the Xen privileged domain – and uses TrouSerS – an open source TCG Software Stack (TSS) [117] – to access the TPM capabilities. It can be decomposed in: (1) server-side stub, the endpoint for the com-

[16]Because the root account in Domain0 has full access to all resources and services, protection against a malicious administrator is not possible in our PoC.

munication with the $TLS\ frontend$ via TCP-based protocol, (2) the Trusted Platform Agent (TPA)[17] in charge of dealing with the life-cycle management of all credentials and implementing a minimal Storage Manager and (3) a software module partially implementing a Trust Manager[18] which performs all cryptographic operations during the handshake and the TLS session.

9.8 Security Considerations

In this section we carry out a short reevaluation of the security requirements adopted from [10] and presented in Section 9.2.1 with regard to our new approach.

(SR1) Secure channel properties: TLS provides secure channel properties during transmission. On the endpoints the TCB offers those properties. Confidentiality and integrity are provided by trusted initialization, isolation of the TCB and platform monitoring. The concept is even able to provide protection against a malicious administrator, because the measurements of the TCB cannot be faked, and if the TCB is properly configured (expressed by its measurements), it should not be possible to tamper with the TCB while running. The TCB also takes care of authenticity and freshness by securely storing nonces and session keys. As a result of platform monitoring, every manipulation of a compartment is noticed and access to sensitive data can be barred if necessary to ensure the security properties. Furthermore, SM provides storage that can preserve secure channel properties in case that data is stored persistently.

(SR2) Authentic linkage of configuration information and secure channel: Authenticity of communication is guaranteed by providing $cert_{TCLS}$ that is used to authenticate the endpoints (cf. Section 9.4). The secure linkage of configuration information to the endpoints is verified by evaluating the $SKAE$, Sig_{BSEK} and Sig_{aD}.

We assume a secure as well as specification conformant creation of $K_{storage}$ (Storage Root Key) and AIK. We further assume that a TCB whose configuration has been evaluated by the counterpart is able to reliably transfer configuration information related to the client compartments and takes care of the secure storage and application of the keys used within the handshake. A possibility for the retrieval of CA keys for verification of signatures is also anticipated.

After a successful evaluation of the credentials transferred, these statements can be made: All keys are bound to the same TCB. This TCB is specified by measurement values incorporated in $SKAE$. Thus, the $properties$ and the $changelog$ sent originate from this TCB because SK_{sign} and SK_{enc} are sealed to this TCB and signed by SK_{SKAE}.

SeK, K_{sign} and K_{enc} are kept inside the TCB during the whole session. Due to the isolation property of the TCB those keys cannot be disclosed to compartments or to other platforms. This is the reason why we claim to provide better forward secrecy for session keys in case of using RSA as key exchange method. Relay attacks as well as attacks to obtain any keys establishing the Trusted Channel are

[17] TPA has been developed within the OpenTC [86] project
[18] we chose to use the OpenSSL library `libcrypto`

not feasible assumed that no hardware attacks are applied.

(SR3) Privacy: With regard to configuration data transmitted, we provide a possibility to send it encrypted to protect potentially sensitive data (see Section 9.5.2). Only the configuration of TCB and the TLS client compartments is reported to the peer, keeping the information disclosed to the other platform as minimal as possible. Furthermore, every communication partner can assess the trustworthiness of its counterpart and thus, make a judgment on whether it will treat personal information according to its security policy.

9.9 Functional Considerations

To meet the functional requirements (cf. Section 9.2.2) following measures have been taken:

(FR1) Fast deployment support: To make a fast and widespread deployment of our approach possible, we decided to adapt TLS as a commonly used protocol to support the exchange of endpoint configuration information. Furthermore, we took an existing TLS implementation (OpenSSL) and adapted it to our needs. Thus, the effort that had to be put into the implementation of our approach was moderate and will presumably be even smaller for other existing TLS implementations, since we had to add features like, e.g, extension support, that normally should already be integrated. Additionally, we only applied mechanisms and concepts already defined in existing specifications. Thus, it is not necessary to go through a time-consuming specification process. Finally, our new concept is able to support all common key exchange methods of TLS.

(FR2) Minimal costs: For the implementation of our concept we used commercial off-the-shelf hardware. Thus, no expensive cryptographic hardware is necessary. Only Open Source Software was used for the realization of the software part. The resulting code is available without charge and incurs no additional license costs for the user. In contrast to [10], we also decoupled TLS certificate keys from a fixed platform configuration by introducing a TCB update mechanism (cf. Section 9.6.2) that enables updates of the TCB without losing track of the states the platform went through.

(FR3) Minimal overhead during handshake: To be able to ensure fast response times, we do not rely on the direct usage of the TPM. This would induce too much overhead since the TPM is currently connected to a LPC-Bus that has only limited bandwidth and its processing power is also very restricted. Thus, we decided to involve the TPM only for system initialization. From then on all functionalities are provided by Trusted Services. Thus, even in a server environment fast response times are ensured. The overhead induced by transferring attestation information is minimal, since the amount of data transferred is comparatively small. We tested this in a testbed and the results showed that there is no significant performance overhead induced by inserting and transferring the extensions.

(FR4) Flexible configuration reporting: By incorporating the possibility to transfer whatever integrity/ configuration information (e.g., see [96]) one may want to pro-

vide, using a black box *properties* field, we ensure interoperability and safeguard forward compatibility of our concept. We showed that, e.g., the TCG approach to use the `TPM_Quote()` [119, p.160] for attestation is also supported (cf. Section 9.4).

(FR5) Backward compatibility: To be sure that also peers only supporting common TLS secure channels can communicate with systems that use our concept, we only used or defined extensions to the different specifications. Those extensions are ignored if they are not supported. Furthermore, we kept the whole implementation separated from the application layer offering a transparent usage of the Trusted Channel. Thus, applications do not have to be adapted to make use of our concept.

9.10 Summary

In this paper we presented a security architecture as well as an adaptation of the TLS handshake to provide a Trusted Channel that combines the security features of a secure channel with the ability to determine the trustworthiness of the communication endpoints.

After a detailed description of our design and its implementation, we showed that our approach is able to meet the strict requirements set in the beginning. By meeting these requirements we are able to provide a means to fight off many threats to today's and tomorrow's distributed applications with a concept that is deployable in the short-term.

In a next step we plan to hand in a RFC for our extensions to the IETF consensus process and we work on adapting IPSec to be used as Trusted Channel protocol. A formal security analysis of the presented TLS handshake is subject to future work as well as the adaptation of other protocols, e.g., SSH, to fit the needs of a Trusted Channel.

Chapter 10

Towards Dependability and Attack Resistance

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10.1 Introduction to Dependable Virtualization

In this chapter, we explore opportunities for dependability and security made available by virtualization, and provide detailed information on how virtualization affects system reliability. We make four contributions: (1) a survey of dependability and security enhancements enabled by virtualization, (2) a prototype demonstrating the effectiveness of hypervisor-based intrusion detection, (3) reliability models and analysis of the effects of virtualization, and (4) an architecture for a reliability-enhanced Xen VMM that leverages a subset of the enhancements.

We describe ways of leveraging virtualization for dependability and security enhancements, such as response to load-induced failures, administration of patches in an availability-preserving manner, enforcement of fail-safe behavior, proactive software rejuvenation, and intrusion detection and protection. We describe in detail a Xen-based implementation of a subset of these enhancements, particularly, intrusion detection and protection. The intrusion detector, called X-Spy, uses a privileged Xen VM to monitor and analyze the complete state of other VMs co-located on the same physical platform. X-Spy is close enough to the target monitored to have a high degree of visibility into the innards of the target (like host-based intrusion detection schemes). At the same time, thanks to the isolation provided by the VMM, X-Spy is far enough from the target to be unaffected even if the target becomes compromised (like network-based intrusion detection schemes). A key challenge in implementing X-Spy was the *semantic gap*, i.e., the proper interpretation of process information gathered from the VMs monitored in a completely different VM.

We provide detailed information on how virtualization affects an important dependability attribute, namely reliability. The VMM is increasingly seen as a convenient layer for implementing many services such as networking and security [41] that were traditionally provided by the operating system. We show why such designs should be viewed with more caution. We use combinatorial modeling to analyze multiple design choices when a single physical server is used to host multiple virtual servers and to

quantify the reliability impact of virtualization. In light of the prevailing trend to shift services out of the guest OS into the virtualization layer, we show that this shift, if not done carefully, could adversely affect system reliability.

We describe a reliability-enhanced Xen VMM architecture, called *R-Xen*, that combines replication, intrusion detection, and rejuvenation. Normally, the Xen VMM consists of a relatively small hypervisor core and a full-fledged privileged VM called *Dom0* that runs a guest OS (Linux). Regular VMs running on the Xen VMM are called *user domains* or *DomUs*. Because of its size and complexity, Dom0 is the weak point in the reliability of the Xen VMM. R-Xen focuses on improving Dom0 reliability (and thereby improving the Xen VMM reliability) through three-fold replication. The three Dom0 replicas each contain X-Spy implementations to mutually monitor each other and thus detect the presence of faults and/or intrusions in the other two. If two replicas report to the hypervisor that the third is corrupted, the hypervisor terminates and rejuvenates the corrupt replica. If the replica terminated happens to be the *primary* replica that provides device virtualization for user domains, then one of the two backups becomes the new primary.

The remainder of the chapter is organized as follows. Section 1.4 describes related work in the area of virtualization-based dependability and virtualization-based intrusion detection. In Section 10.2, we describe at a high-level several dependability and security enhancements (including intrusion detection and protection) that are made possible by virtualization. Section 10.3 describes X-Spy, our Xen-based prototype implementation of intrusion detection and protection. Section 10.4 analysis the reliability impact of virtualization and highlights the importance of VMM reliability to the overall reliability of a virtualized physical node. Motivated by the conclusions of our reliability analysis and leveraging our X-Spy implementation, Section 10.5 describes an architecture for a more reliable Xen VMM.

10.2 Using Virtualization for Dependability and Security

Commodity operating systems provide a level of dependability and security that is much lower than what is desired. This situation has not changed much in the past decade. Hence, the focus has shifted to designing dependable and secure systems around the OS problems. Thanks to the flexible manner in which VM state can be manipulated, virtualization can enable such designs. In particular, VM state, much like files, can be read, copied, modified, saved, migrated, and restored [41]. In this section, we give several examples of dependability and security enhancements made possible by virtualization.

Coping with Load-Induced Failures: Deploying services on VMs instead of physical machines enables a higher and more flexible resilience to load-induced failures without requiring additional hardware. Under load conditions, the VMs can be seamlessly migrated (using live migration [23]) to a lightly loaded or a more powerful physical machine. VM creation is simple and cheap, much like copying a file. In response to high-load conditions, it is much easier to dynamically provision additional VMs on under-utilized physical machines than to provision additional physical machines. This flexibility usually compensates for the additional resources (mainly memory) needed by the hypervisor.

Patch Application for High-Availability Services: Typically, patch application involves a system restart, and thus negatively affects service availability. Consider a service running inside a VM. Virtualization provides a way of removing faults and vulnerabilities at run-time without affecting system availability. For this purpose, a copy of the VM is instantiated, and the patch (be it OS-level or service-level) is applied on the copy rather than on the original VM. Then, the copy is restarted for the patch to take effect, after which the original VM is gracefully shut down and future service requests are directed to the copy VM. To ensure that there are no undesirable side effects due to the patch application, the copy VM may be placed under special watch for a sufficiently long time while its post-patch behavior is being observed before the original VM is shut down. If the service running inside the VM is stateful, then additional techniques based on a combination of VM checkpointing (e.g., [7]) and VM live migration [23] may be used to retain network connections of the original VM and to bring the copy up-to-date with the last correct checkpoint.

Enforcing Fail-Safe Behavior and Virtual Patches: The average time between the point in time when a vulnerability is made public and a patch is available is still measured in months. In 2005, Microsoft took an average time of 134.5 days for issuing critical patches for Windows security problems reported to the company [126]. Developing patches for a software component is a time-consuming process because of the need to ensure that the patch does not introduce new flaws or affect the dependencies between the component involved and other components in the system. In many cases, a service administrator simply does not have the luxury of suspending a service immediately after a critical flaw (in the OS running the service or the service itself) becomes publicized until the patch becomes available.

Virtualization can be used to prolong the availability of the service as much as possible while at the same time ensuring that the service is fail-safe. We leverage the observation that publicizing a flaw is usually accompanied by details of possible attacks exploiting the flaw and/or symptoms of an exploited flaw. Developing an external monitor to identify attack signatures or symptoms of an exploited flaw may be done independently of patch development. The monitor may also be developed much faster than the patch itself, because the monitor may not be subject to the same stringent testing and validation requirements.

Consider a service running inside a VM rather than directly on a physical machine. Then, a VM-external monitor, running in parallel to the VM, can be used to watch for these attack signatures or detect the symptoms of exploitation of the flaw. If attack signatures are known, the VM-external monitor can be used to block the attack, e.g. by filtering the incoming network stream, to terminate interaction with the attack source, or to protect targeted structures inside the VM, e.g. the system call table. If only symptoms of exploitation are known, detection of a compromise can be used to immediately halt the VM. The monitor could be implemented at the VMM level or in a privileged VM (such as Dom0 in Xen [11]). If it is important to revert the service to its last correct state when a patch becomes available, then the above technique can be augmented with a checkpointing mechanism that periodically checkpoints the state of the service with respect to the VM (e.g., [7]).

Proactive Software Rejuvenation: Rebooting a machine is an easy way of rejuvenating software. The downside of machine reboot is that the service is unavailable during the reboot process. The VMM is a convenient layer for introducing hooks to proactively rejuvenate the guest OS and services running inside a VM in a performance- and availability-preserving way [93]. Periodically, the VMM can be made to instantiate a *reincarnation VM* from a clean VM image. The booting of the reincarnation VM is

done while the original VM continues regular operation, thereby maintaining service availability. One can view this technique as a generalization of the proactive recovery technique for fault-tolerant replication proposed by Reiser and Kapitza [93].

As mentioned above in the context of patch application, techniques based on VM checkpointing and live migration may be used to seamlessly transfer network connections and the service state of the original VM to the reincarnation VM. It is possible to adjust the performance impact of the rejuvenation procedure on the original VM's performance. To lower the impact, the VMM can restrict the amount of resources devoted to the booting of a reincarnation VM and compensate for the restriction in resources by allowing more time for the reboot to complete.

One can view the above type of rejuvenation as a *memory-scrubbing* technique for reclaiming leaked memory and recovering from memory errors of the original VM. More importantly, such a periodic rejuvenation offers a way to proactively recover from errors without requiring failure detection mechanisms (which are often unreliable) to trigger the recovery.

Intrusion Detection and Response: Based on the location of the intrusion detection sensors, intrusion detection system (IDS) implementations are broadly classified into host-based IDS (HIDS) and network-based IDS (NIDS) [46]. A NIDS monitors network traffic from and to the target, and analyzes the individual packets for signs of intrusion. Because of its isolation from the target monitored, a NIDS decreases its susceptibility to attacks and is largely unaffected by a compromised target. However, as network traffic becomes increasingly encrypted and as the NIDS has no direct knowledge of the effects or properties of the attack targets, the usefulness of NIDS is decreased. The fact that not all intrusions may manifest their effects in the form of malicious traffic also lowers the utility of NIDS. The sensors of a HIDS are placed on the target machine itself, giving them a high degree of visibility into the internals of the target, enabling closer monitoring and analysis of the target than NIDS does. However, the location of HIDS on the same "trust compartment" as the target is also a disadvantage: after an intrusion into the target, the HIDS may no longer be trusted.

Virtualization provides a way of removing the disadvantages of HIDS and NIDS, while retaining their advantages. In our approach, the sensors are placed in a special privileged VM (called the *secure service VM* or SSVM) used for monitoring other VMs hosting regular production services (called *production VMs* or PVMs). The placement of the sensors on the same physical machine but in a different VM allows monitoring and analysis of the complete state of other VMs via the VMM, and at the same time, keeps the sensor out of reach of a potentially compromised VM and in a secure vantage position.

The twin characteristics of proximity to the target and isolation from the target also make the SSVM a convenient location for implementing intrusion response mechanisms. The secure vantage point of the SSVM allows one to implement otherwise difficult responses, e.g., even a simple response like 'shutdown a compromised system' may not be effectively triggered from inside the compromised system. On the other hand, it is easy and effective to suspend the operations of a compromised PVM from the SSVM. In addition, the SSVM can instruct the VMM to provision a healthy replacement PVM or block suspicious system calls that may potentially tamper with the integrity of the kernel.

For effective rejuvenation of a compromised PVM by re-provisioning a new PVM, it is not sufficient to merely boot the new PVM from a clean state. The new PVM might still possess all the vulnerabilities of the compromised one. Hence, it is important to perform a forensic analysis of the compromised PVM's state to remove as many

vulnerabilities as possible. Such an analysis is facilitated by the virtualized environment hosting the SSVM. The SSVM can obtain not only modified files of a suspended PVM, but also its complete run-time state from the memory dump created at the time of suspension. The memory dump can be examined using the same techniques as the one used to observe the state of a running PVM from the SSVM for the purpose of intrusion detection.

10.3 Xen-based Implementation of Intrusion Detection and Protection

In this section, we describe the prototype implementation of a subset of the security enhancements mentioned above, namely, intrusion detection and protection for VMs. Later, in Section 10.5, we leverage the implementation for enforcing fail-safe behavior and for triggering software rejuvenation in our construction of R-Xen.

10.3.1 Intrusion Detection and Protection for Xen Virtual Machines

We have implemented an intrusion detection and protection framework called *X-Spy*. The core idea is to use a secure service VM (SSVM) that monitors one or more production VMs (PVM). The SSVM performs the following functions:

Lie Detection The SSVM accesses the memory of the PVM and compares actual critical system data (processes, mounts, etc.) against data obtained by executing normal Unix commands inside the PVM. If the comparison yields discrepancies, then that is indicative of a compromised PVM. In contrast to earlier hypervisor-based intrusion detection work, X-Spy's detection mechanisms are more comprehensive and include lie detection at the level of processes, network connections, modules, and file system mounts.

Protection We have added a system call inspector to Xen that allows the monitoring of the system calls within the PVM for the purpose of protecting relevant forensic information (like log files) and the integrity of the kernel (kernel structures, modules, and memory).

X-Spy uses the Xen [11] VMM developed by Cambridge University and guest VMs running the Linux 2.6 operating system. Nevertheless, the concepts such as system call analysis and lie detection can be applied to other operating systems such as Microsoft Windows. All X-Spy components are implemented either in the Xen hypervisor or in the SSVM. While their implementation logic depends on the guest OS, X-Spy does not require any modification to the guest OS of the PVM.

Limitations

To overcome the semantic gap, we assume some knowledge of the kernel structures of the guest operating system (specifically, Linux kernel 2.6) so that X-Spy components

can be appropriately coded. If the guest operating system is upgraded to a newer version in which kernel structures are different, then the X-Spy components need to be re-coded appropriately. That fact may be an impediment to commercializing X-Spy, as it implies an ongoing commitment to develop and patch X-Spy components to keep pace with upgrades to the guest operating system.

For detecting hidden processes, X-Spy requires that the scheduler of the PVM's guest OS keep a list of processes that need to be scheduled in a standard place within a known memory structure. If an attacker is able to replace the scheduler with her own one having a different list of processes, the detection approach would be subverted. That is why it is important to protect the integrity of the kernel code (for example, using mechanisms that we describe in Section 10.3.4).

The SSVM needs read access to the memory of the PVMs for the purpose of monitoring them. In addition, it must be possible to do an SSH login to the PVM from the SSVM and execute normal Unix commands. These requirements are contrary to the isolation guarantees of the hypervisor. The SSVM itself could become a high-value attack target, and accordingly, needs stronger protection. Several measures can be taken to strongly reduce the potential of the SSVM getting compromised. For example, as the SSVM is a special-purpose VM (in contrast to PVMs), it can be hardened, its functionality reduced solely to that of monitoring the PVMs, and its access restricted through a specific administrative interface.

10.3.2 Architecture of X-Spy

The architecture of the X-Spy intrusion detection framework is shown in Figure 10.1. Our architecture consists of a PVM and a SSVM running on top of the same hardware and Xen hypervisor. In our implementation, both the SSVM and the PVM run Linux kernel 2.6. The SSVM obtains the run-time state of the PVM through the Xen hypervisor, which is at a lower level of abstraction than both the SSVM and the PVM. The SSVM has access to the raw devices of the PVMs (memory, disk, network); however, the difficulty lies in the SSVM properly interpreting the data because of a semantic gap [21]. For example, the physical memory of the host system will be made available in chunks as *pseudo-physical* memory to the VMs. In addition, the (possibly different) operating systems of the VMs use a virtual address space on top of the physical memory, leading to the problem of properly interpreting raw memory locations in a different context.

10.3.3 Intrusion Detection by means of a Lie-Detector

The basic idea of the Lie-Detector is to compare the insider and outsider views of the system to identify objects (processes, files etc.) that try to hide themselves from the operating system [12]. Such behavior is typical of *rootkits*, which are then used to hide other (typically malicious) software, but is also sometimes characteristic of DRM functionality (e.g. the XCP content protection technology by Sony BMG in 2005). The Lie-Detector (Figure 10.1) consists of three major functionalities:

1. PVM Information Collection: The Lie-Detector collects information about the PVM by two different means: the *native* interface and the *frontDoor* interface.

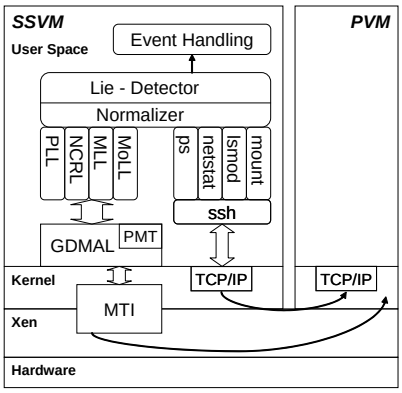


Figure 10.1: Architecture of the X-Spy Lie-Detector components.

2. PVM Information Normalization: The PVM information collected via the native interface is normalized to a format equivalent to that of commands executed through the frontDoor interface.

3. Analyze-and-Compare: The normalized information from the native and *front-Door* interfaces is then compared to identify differences that are indicative of maliciously hidden system resources and to minimize false positives. Any findings will be reported through the *Event Handling* component.

We describe the above functionalities in detail below.

Memory Translation Interface (MTI) One of the main components of X-Spy is a Memory Translation Interface (MTI) that allows the SSVM full access to a PVM's pseudo-physical and virtual memory in a convenient fashion. The MTI has two parts:

1. An extension to the Xen hypervisor, which performs address translation and traversal of the page tables.

2. A Linux kernel device driver that runs in the SSVM kernel and provides two interfaces, namely, `/dev/mem_domX` and `/dev/kmem_domX`. These interfaces are functionally equivalent to the `/dev/mem` and the `/dev/kmem` device files, respectively, and allow the root user in the SSVM kernel access to the PVM's physical memory and kernel memory contents, respectively.

One challenge to overcome when implementing the MTI was that the SSVM cannot access a PVM's foreign memory as it corresponds to a different context. Therefore, the MTI has to emulate the memory management unit by translating the address to the right format and re-mapping it from the PVM memory space onto the memory space of the Lie-Detector process running in the SSVM. For this purpose, we have developed two user-space libraries that the MTI uses: the *Guest Domain Memory Access Library* or GDMAL and the *Process Memory Translator* or PMT. The GDMAL provides read-write access to the PVM's memory. Within the PVM's memory, the PMT allows access to the virtual address space of PVM processes. In addition, the PMT provides some helper functions to facilitate the use of the `/dev/mem_domX`

and `/dev/kmem_domX` interfaces. The PMT performs the process address translation by extracting the memory management information for the process from the OS-specific task (process) description data structure. When the guest OS is Linux, as in our case, the PMT extracts the `mm_struct` data structure from the `task_struct` data structure.

The *native interface* is used to collect PVM information "from the outside" through the raw access made available by the Xen hypervisor, e.g. by accessing the PVM's memory via the MTI, and to collect host-specific data via special user-space libraries that we have developed, namely the process list library (PLL), the network connection and routing library (NCRL), and the module list library (MLL).

The second interface, called the *frontDoor interface*, is used to obtain PVM information by doing an SSH login to the PVM and executing normal Unix commands. The Lie-Detector normalizes the information collected from both interfaces, and then compares them. If the comparison yields discrepancies in the information collected from those two sources, this is strongly indicative of an intrusion. Obviously, it is not possible to obtain information through the frontDoor interface and the native interface at exactly the same time. This timing difference may lead to false positives, and we explain below how to overcome this problem. We implement comparison methods for processes, network connections, kernel modules, file system mounts and files.

The MTI provides access to the PVM's raw kernel virtual memory but lacks any semantic context. To fix this shortcoming, we manually created a *memory offset file* for each library[1]. Based on these files, the libraries such as PLL, NCRL, and MLL, implement the logic to extract all data values of interest from the raw kernel virtual memory. Each offset file stores the offset values of the start of each data item of interest from the beginning of the containing structure.

Process List The PLL acts on information provided by the MTI to generate output similar to that of the `ps` command. This is done by accessing and then traversing the doubly-linked circular task list via the MTI.

Our comparison is based on the multitude of information extractable from this `task_struct` data structure, such as PID, state, parent, open files, registers, priorities, locks, and memory management information. However, not all fields in the data structure are used. For example, the running time of a process as seen by the native interface query and the frontDoor interface query are bound to slightly differ, owing to the difference in the time of query.

This comparison will detect processes in the PVM that actively try to hide their presence or change their appearance (e.g. the owner) from queries made from within the PVM. This will identify rootkit-like behavior, as non-hiding processes can be identified by more conventional (non-hypervisor-based) malware detection tools.

Note that simply comparing the process information from the native and frontDoor interfaces results in false positives because of frequent changes to the process table. We fix this by executing a native access (outsider view) before and after the frontDoor query (insider view). If a given process disappeared in the second query but is again visible in the third, we consider it to be an intrusion. If it does not reappear, we assume that the process merely terminated.

Network Connections and Routing We obtain information about IPv4 connections, Unix socket connections, and IPv4 routing through the native and frontDoor inter-

[1] With some effort, it is possible to generate the offset files automatically at kernel compilation time.

faces. For the native interface queries, we have developed the NCRL library, which uses the MTI to collect information equivalent to that obtained from three commands: '`netstat -an -inet`' for IPv4 connections, '`netstat -an -unix`' for Unix Socket connections, and '`netstat -rn`' for IPv4 routing. This information can then be used to discover hidden network connections.

Similar to the timing problem in the Lie-Detector comparison of process information, we face a timing problem in the comparison of network connection information because of network connections that were terminated or started in the time interval between the native interface query and the frontDoor interface query. The solution here is again to reduce false positives by using three queries[2].

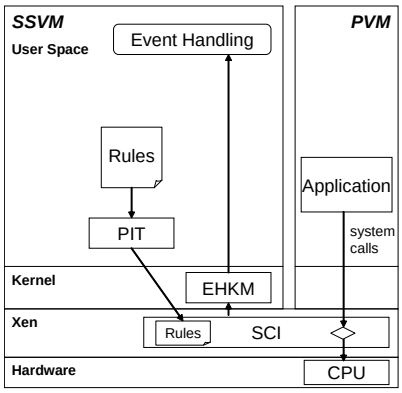
Module List To obtain information about the PVM's kernel modules we have developed another user-space library called the MLL for collecting information from the native interface query. The frontDoor interface query uses the `lsmod` command, which outputs the contents of `/proc/modules` displaying the kernel modules currently loaded. In addition to the native interface and frontDoor interface queries, the MLL also queries a third Xen interface for detecting hidden Linux loadable kernel modules (LKMs). LKMs are a way to link object code without interruption to the Linux kernel while it is running. Such LKMs are automatically registered at loading time, but it is possible for an LKM to un-register itself after loading. In such a case, the LKM can hide even from a native interface query (as the adore-ng rootkit indeed does; see Section 10.3.5). To address this issue, we established a *shadow module list* in the hypervisor. The hypervisor traps the `init_module` system call and analyzes the ELF header section of the object file to get the module name and stores the name in the shadow module list. The hypervisor also traps the `delete_module` system call to remove entries from the shadow module list. As the hypervisor address space cannot be accessed by the PVM, the shadow module list cannot be altered by an intruder. The Xen interface query shows the contents of the shadow module list and is taken as reference for comparison with the results of the native interface and frontDoor interface queries. If the results from the native interface and/or the frontDoor interface queries do not list an entry from the shadow module list, we conclude that the module in question is hidden.

Mounts The frontDoor interface uses the `cat /proc/mounts` command, which provides a list of all mounted file systems in the PVM. An obvious alternative would have been to use the output of the `mount` command; however that alternative is less useful and secure because the command merely outputs the contents of the `/etc/mtab` file, and it is easy to mount a file system without an entry showing up in the `/etc/mtab` file by using the `mount -n` command.

The mount list library (MoLL[3]) operates on the PVM information about mounted file systems collected via the native interface query. The starting symbol for obtaining the information is the `task_struct` structure of the idle task (however, the entry for any task would be adequate), from where the MoLL gains access to the `vfsmnt` circular list. The list provides complete information about all file systems currently mounted.

[2]Note that the frontDoor query is made through an SSH connection, which will show up only in the frontDoor query but in neither of the interface queries.

[3]The MoLL should not be confused with MLL, the module list library.



The mount information gathered from the native interface query is used as the reference against which the information from the frontDoor interface is compared. If there are mounted file systems that appear in the former but not in the latter, we take this as an indication of a hidden malicious process because mount information is relatively static, and hence false positives are not a big concern.

File system In the case of the file system, bridging the semantic gap in general implies the use of raw access to the physical disk and the related traffic to rebuild the file system structures of the guest operating system of the PVM in the context of the SSVM. Accessing file systems mounted by another operating system is feasible even for disparate operating systems, e.g. Microsoft Windows and Linux, as for example shown in [57]. Xen can use a Linux file system existing on the Xen/Dom0 level to boot and launch guest domains. This same file system can then be mounted read-only by the SSVM. We then retrieve the file information via the frontDoor and compare it with the information of the file system mounted by the SSVM.

For efficiency and simplified forensic recording, we use a basic read-only file system and add the CoWNFS *copy-on-write* file system [67, 94]. This allows us to store the changes for multiple runs for later forensic analysis and protects the original state of the system from any (potentially malicious) changes. This combined file system was then used as an NFS mountable file system for booting the PVM.

10.3.4 Protection of System Integrity and Forensic Information by means of System Call Inspection

We now outline how X-Spy's System Call Inspection component is used to protect the system against intrusions.

Protection of Forensic Information

In case of a successful intrusion it is highly desirable to protect as much forensic information as possible. A smart intruder would want to hide all traces of the intrusion, e.g., by altering log-files[4] such as the `wtmp/utmp` and the `/var/log/messages` files. While these files cannot be modified by normal users, intruders with root access

[4]Note that the above protection scheme for log files can easily be extended to protect other important files, such as Xen VM configuration files, through additional rules in the rule set.

can. One way to address this issue would be to use a hardened system (e.g., SELinux). However, this protects only if the superuser is not allowed to change the SELinux rules in a running System.

In a virtualized environment as considered here, we have the possibility to protect important files by intercepting the system call sequence in the PVM through the Xen hypervisor. For this purpose, we added a module to the Xen hypervisor, the *System Call Inspector* (SCI), which can inspect all[5] system calls occurring in the PVM and either block or accept calls depending on a set of *rules*. These rules are stored and edited in the SSVM (and therefore out of reach of any activity in the PVM), and can be loaded into the SCI (in binary form) via the Policy Installation Tool (PIT).

X-Spy implements a functionality for checking and fine-tuning system calls by instrumenting the system call handling chain. An `int 0x80` instruction is intercepted by an interrupt handler located in the Xen context where checks against the previously introduced rule set are done. Only after passing the checks is the call redirected to the PVM Kernel, where the normal system call handler is invoked; otherwise, the call returns without any action being taken. In certain cases, the system call is allowed after some fine-tuning, e.g., a modification of the parameters so that the call conforms to the rule set specified. The amount of performance overhead depends on the type of checks and fine-tuning being done for a particular system call.

As the interception of the system call happens in the Xen context, the problem of semantic gap has to be overcome to determine which system calls actually merit additional checks. For our aim of protecting forensic information, system calls performing file operations are essential. We protect forensic information by preventing calls that rename, link, unlink, or delete log files. Furthermore, we limit access to log files by permitting only the append operation on them. To ensure that a malicious process cannot bypass the checking, we normalized the paths.

If the SCI finds that an application in the PVM tries to initiate a system call that is not allowed according to the rule set, it will block or modify it and send a corresponding event through an *event handling kernel module* (EHKM) in the SSVM to the high-level event handling component with information about the violated rule and the corresponding process in the PVM.

Protection of Binaries against User-Space Rootkits

The mechanism used for protecting forensic information can also be used to protect binaries from being altered by an intruder. Many user-space rootkits try to alter `ps` or `netstat` to hide their presence or to install a back door by modifying the `openssh` binary. While earlier tools, such as Tripwire, can *detect* the alteration of a binary or a library, our event-driven approach to check system calls and their arguments can actually *prevent* their alteration.

In addition, it is possible to restrict read/write access to an executable, but still allow its execution. Based on the corresponding rule set, the module we have implemented in the Xen hypervisor checks whether a system call is trying to change, delete, link, or rename a binary, and if so, the call is denied. As execution of a binary normally happens through the `execve` system call without actually opening the binary file, it is even possible to add a rule that forbids the opening of certain binaries completely without disallowing their execution.

[5]Note that Xen implements a “fast trap” mechanism to enhance performance. If Xen calls are to be monitored as well, then this mechanism need to be disabled.

Kernel Sealing

X-Spy also implements *kernel sealing*, a well-known method to protect a system or prevent intrusions. The kernel memory can be accessed directly by reading or, more dangerously, by writing to the `/dev/mem` or `/dev/kmem` device files. The rule set of the X-Spy event-driven module in the Xen hypervisor was updated to restrict access to those files, so that only read requests are allowed and write requests return an error result without performing the write operation.

Accessing the kernel memory by loading a kernel module or writing directly to /dev/(k)mem is potentially dangerous because it allows an intruder to establish its own interface to the kernel; thereafter, the intruder can easily place malicious code in the kernel and have full access to the file system and other kernel internals. X-Spy uses a technique called *white-listing* by which all kernel modules allowed to be loaded are explicitly specified along with their respective SHA-1 hash values. If the module to be loaded at run-time is not specified in the white-list or if it has an incorrect hash value, X-Spy prevents the module from being loaded by preventing the system call from reaching the PVM kernel space. Note that our X-Spy implementation does not offer protection against buffer overflows on systems calls.

Pre-Checking of Binaries

An effective way of protecting a PVM from user-space rootkits or other malicious software is to check the hash of every binary, prior to its execution, against a white-list of pre-calculated hashes and to allow its execution only if there is a match. Computing the hash of the binary has to be done out of the reach of a potential intruder in the PVM and should also not require modification of the PVM's OS. To meet these conditions, X-Spy computes the hash of the binary in the SSVM. To enable such a computation, it is necessary that the SSVM has all partitions of the PVM mounted; furthermore, the binary should not be on a RAM disk, on network file system, or on an encrypted file system that the SSVM cannot access. An alternative would be to do the computation in the hypervisor, which would require overcoming the semantic gap problem.

For computing the hash of the binary in the SSVM, we use a technique called *memory scanning*, which involves loading the complete `.text` and `.data` sections of an ELF binary into memory by setting the program counter to the next page, asking the PVM kernel to load the page, and then hashing it while handling the page fault.

If the hash cannot be verified the hypervisor invalidates all of the memory and returns the control back to the guest domain. Because of the invalid `.text` section to which the PC points, the process will crash. Note that relying on support from the PVM's guest OS does not necessarily constitute a security gap, because a non-cooperative PVM kernel would lead to a wrong hash value and in this case, as seen above, the process will be forced to quit.

10.3.5 Evaluation

While the conceptual limitations have been discussed in Section 10.3.1, we will now evaluate the effectiveness and efficiency of our implementation.

Functional Evaluation

The goal of the Lie-Detector is to expose malicious code. Once all information is visible, we no longer focus on whether a particular LKM or process is hostile.

We first evaluated our implementation using a set of functional tests that implement typical rootkit functionality, such as hiding processes, kernel modules, mounts, or network connections. Our Lie-Detector correctly exposed all attempts to hide such critical information.

We then tested the system using the adore-ng rootkit [107]. Adore-ng is an LKM-based rootkit for Linux kernels which allows one to hide files and directories, processes, and network connections. The basic rootkit consists of the kernel module and a user-space control program. While currently this is the only widely available rootkit for the Linux kernel 2.6, other similar rootkits exist for earlier kernel versions and may be ported to the 2.6 kernel.

In a fully protected X-Spy system, the rootkit cannot even be installed as the insertion of modules is restricted through the SCI interception of the respective system calls and the white-listing of allowed modules. After explicitly allowing the rootkit to insert itself into the kernel, we used its control program to hide processes, files, and network connections. The X-Spy Lie-Detector component reported these hidden resources faithfully by comparing the responses from the native and frontDoor interfaces as described in Section 10.3.3. Although the adore-ng kernel module will remove itself from the list of modules visible with `lsmod`, detection of the module by the Lie-Detector is possible with the help of the shadow module list (see discussion in Section 10.3.4).

The rule set used in X-Spy's event-driven protection mechanism contained about 110 rules, e.g. to protect forensically relevant files (e.g. `/var/log/messages` and `/var/log/wtmp`) and to prevent access to raw memory (`/dev/(k)mem`), security relevant configuration files (`/etc/ssh.config`), and operating system tools (`/bin/ls`). In addition, we specified an explicit list of allowed kernel modules (module white-listing). Once the rule set was active, it either generated security events with information about the offending processes in the PVM or successfully prevented the deletion or truncation of log-files and the modification of configuration and utility files.

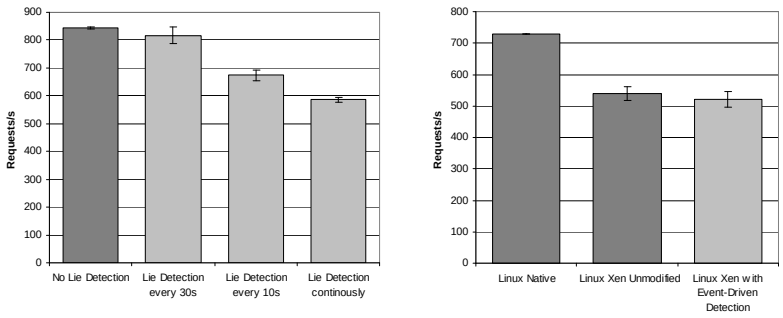
Performance Impact

To measure the performance impact of the Lie-Detector and the event-driven approach, we used a single machine implementing a web server scenario. The PVM hosted an Apache web server, and multiple clients were simulated using the *ab* performance benchmarking tool (see `http://httpd.apache.org/docs/2.0/programs/ab.html`). The networks were virtual and internal to this machine.

Figure 10.2(a) shows that the performance impact of the Lie-Detector depends on how often it is run. The overhead is roughly 31% when it is running continuously, 20% when it is run every 10 sec, and 4% when it is run every 30 sec.

Most practical applications will run infrequent scans. In this case, the performance impact of X-Spy is negligible, particularly when compared with the performance reduction of moving Linux into a VM.

In a real-world setting, the frequency of "Lie Detection" should be chosen based on the expected time until an intrusion occurs and the expected time until such an intrusion is detected. The latter is an important factor because it denotes the critical time window between the intrusion and its detection when the PVM is at the mercy of the intruder, who can take arbitrary actions (such as installing a fake website or copying private information onto a different system). If the PVM runs a critical service in which the critical time window should be minimized, then the Lie-Detector should be run continuously.



(a) Impact of Lie-Detector

(b) Impact of protective measures

Figure 10.2: Performance impact of X-Spy components: number of fulfilled requests per second in the HTTP benchmark.

As seen in Figure 10.2(b), the event-driven method results in a performance loss of about 4%. Compared with the 34% overhead incurred by changing from a service running on a non-virtualized platform to that running on a Xen-based PVM, the loss incurred by the event-driven approach is minor.

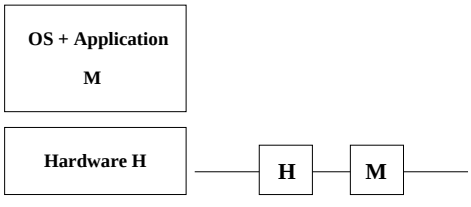
10.4 Quantifying the Impact of Virtualization on Node Reliability

In this section, we use combinatorial modeling to perform a reliability analysis of redundant fault-tolerant designs involving virtualization on a single physical node and compare them with the non-virtualized case. The results of the analysis highlight the importance of improving the reliability of the hypervisor.

We consider a model in which multiple VMs run concurrently on the same node and offer identical service. We derive lower bounds on the VMM reliability and the number of VMs required for the virtualized node in order to have better reliability than in the non-virtualized case. We also analyze the reliability impact of moving a functionality common to all VMs out of the VMs and into the VMM. In addition, we analyze the reliability of a redundant execution scheme that can tolerate the corruption of one out of three VMs running on the same physical host, and compare it with the non-virtualized case. Our results point to the need for careful modeling and analysis before a design based on virtualization is used.

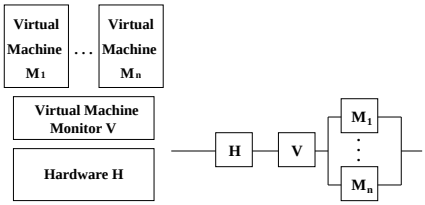
Combinatorial modeling and Markov modeling are the two main methods used for reliability assessment of fault-tolerant designs [58]. We chose combinatorial modeling because its simplicity enables easy elimination of "hopeless" choices in the early stage of the design process. In combinatorial modeling, a system consists of series and parallel combinations of modules. The assumption is that module failures are independent. In a real-world setting, where module failures may not be independent, the reliability value obtained using combinatorial modeling should be taken as an upper bound on the system reliability.

Non-Virtualized (NV) Node: For our reliability assessment, we consider a non-virtualized single physical node as the base case. We model the node using two mod-



(a) Architecture (b) Combinatorial Model

Figure 10.3: Non-virtualized node



(a) Architecture (b) Combinatorial Model

Figure 10.4: Node with n VMs

ules: hardware (H) and the software machine (M) consisting of the operating system, middleware, and applications (Figure 10.3(a)) . Thus, the node is a simple serial system consisting of H and M, whose reliability is given by $R_{sys}^{NV} = R_H R_M$, where R_X denotes the reliability of module X (Figure 10.3(b)).

Virtualized Node with n Independent, Identical VMs: Figure 10.4(a) shows a physical node consisting of H, a type-1 VMM (V) that runs directly on the hardware (such a VMM is referred to as a *hypervisor*), and one or more VMs ($\{M_i\}, i \geq 1$). The VMs provide identical service concurrently and independently (i.e., without the need for strong synchronization). For example, each VM could be a virtual server answering client requests for static web content. Thus, the node is a series-parallel system (Figure 10.4(b)) whose overall reliability is given by $R_{sys}^{n} = R_H R_V[1 - \prod_{i=0}^{n}(1 - R_{M_i})]$. Here, we consider the reliability of the hardware to be the same as that in the non-virtualized case because the underlying hardware is the same in both cases. An obvious concern is whether the hardware in the virtualized node will register a significant drop in reliability due to load/stress compared with the non-virtualized node. However, this concern does not apply to our context of application servers in a data center, in which typical hardware utilization in a non-virtualized node is abysmally low (less than 5%) and n is typically in the low tens of VMs.

The condition for the n-replicated service to be more reliable than the non-virtualized service is given by $R_{sys}^{n} > R_{sys}^{NV}$. i.e., $R_H R_V[1 - \prod_{i=0}^{n}(1 - R_{M_i})] > R_H R_M$. For simplicity, let $R_{M_i} = R_M$ for all $1 \leq i \leq n$. This is a reasonable assumption, as each VM has the same functionality as the software machine M in the non-virtualized case. Then, the above condition becomes

$$R_V[1 - (1 - R_M)^n] > R_M. \tag{10.1}$$

Inequality (10.1) immediately yields two conclusions. First, if $n = 1$, then again the above condition does not hold ($R_V < 1$). What this means is that it is necessary to have some additional coordination mechanism or protocol built into the system to compensate for the reliability lost by the introduction of the hypervisor. In the absence of such a mechanism/protocol, simply adding a hypervisor layer to a node will only decrease node reliability. Second, if $R_V = R_M$, then it is obvious that above condition does not hold.

It is clear that the *hypervisor has to be more reliable than the individual VMs*. The interesting question is how much more reliable. Figure 10.5 shows that for a fixed R_M value, the hypervisor has to be more reliable when deploying fewer VMs. The graph also shows that, for fixed values of R_M and R_V, there exists a lower bound on n below which the virtualized node reliability will definitely be lower than that of a non-virtualized node. For example, when R_M = 0.1 and R_V = 0.3, deploying fewer

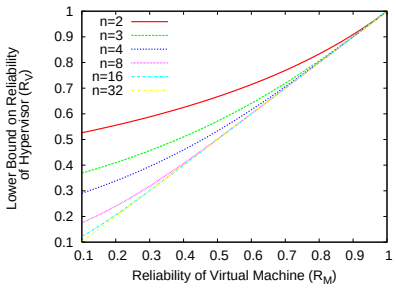


Figure 10.5: Lower bound on the hypervisor reliability for a physical node with n independent and concurrently operating VMs providing identical service.

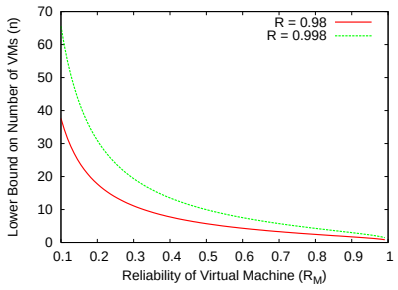


Figure 10.6: Lower bound on the number of VMs to achieve desired reliability R for a physical node with n independent and concurrently operating VMs providing identical service when $R_V = 0.999$.

than 4 VMs would only lower the node reliability. This is a useful result, as in many practical settings, R_M and R_V values may be fixed, e.g., when the hypervisor, guest OS, and application are commercial off-the-shelf (COTS) components with no source-code access.

The equation for R^n_{sys} also suggests that by increasing the number of VMs, the node reliability can be made as close to the hypervisor reliability as desired. Suppose we desire the node reliability to be R, where $R < R_V$. Then, $R = R_H R_V[1-(1-R_M)^n]$. Assume that the hardware is highly reliable, i.e., $R_H \simeq 1$. Then, the above equation becomes the inequality,

$R < R_V[1 - (1 - R_M)^n]$
$\Longrightarrow (1 - R_M)^n < 1 - \frac{R}{R_V}$
$\Longrightarrow n.\log(1 - R_M) < \log(1 - \frac{R}{R_V})$

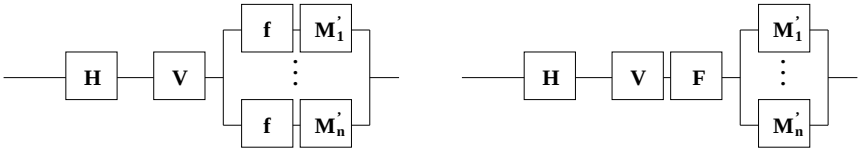
Dividing by $\log(1 - R_M)$, a negative number, we obtain,

$$n > \frac{\log(1 - \frac{R}{R_V})}{\log(1 - R_M)}. \tag{10.2}$$

Inequality (10.2) gives a lower bound on the number of VMs required for a virtualized physical node to meet a given reliability requirement. In practice, the number of VMs that can be hosted on a physical node is ultimately limited by the resources available on that node. Comparing the lower bound with the number of VMs that can possibly be co-hosted provides an easy way of eliminating certain choices early in the design process.

Figure 10.6 shows the lower bound for n for two different R values (0.98 and 0.998) as the VM reliability (R_M) is increased from roughly 0.1 to 1.0, with the hypervisor reliability fixed at 0.999. The figure shows that for fixed R_V and R_M values, a higher system reliability (up to R_V) can be obtained by increasing the number of VMs hosted. However, when n is large, one is faced with the practical difficulty of obtaining sufficient diversity to ensure that VM failures are independent.

Moving Functionality out of the VMs into the Hypervisor: We now analyze the reliability impact of moving a functionality out of the VMs and into the hypervisor. As before, our system model is one in which a physical node has $n \geq 1$ independent



(a) Configuration C_1: Functionality f implemented within each VM

(b) Configuration C_2: Functionality F implemented as part of Hypervisor

Figure 10.7: Moving functionality out of the VMs into the hypervisor

and concurrently operating VMs providing identical service. Consider a functionality f implemented inside each VM. Then, each VM M_i can be divided into two components, f and $M_i^{'}$, the latter representing the rest of M_i. Figure 10.7(a) shows the reliability model for a node containing n such VMs. Let us call this node configuration C_1. Further, suppose that the functionality f is moved out of the VMs and substituted by component F implemented as part of the hypervisor. Now, the new hypervisor consists of two components F and the old hypervisor V. Figure 10.7(b) shows the reliability model for a node with the modified hypervisor. Let us call this node configuration C_2.

We now derive the condition for C_2 to be at least as reliable as C_1. For simplicity, let us assume that $R_{M_i'} = R_{M'}$ for all $1 \leq i \leq n$. Then, the desired condition is $R_{sys}^{C_2} \geq R_{sys}^{C_1}$

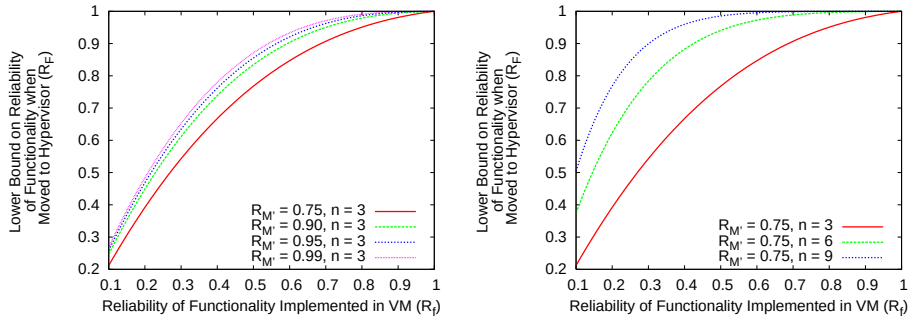
$$\implies R_H R_V R_F[1-(1-R_{M'})^n] \geq R_H R_V[1-(1-R_f R_{M'})^n]$$

$$\implies R_F \geq \frac{[1-(1-R_f R_{M'})^n]}{[1-(1-R_{M'})^n]}. \tag{10.3}$$

It is easy to see from Figure 10.7 that if there is only one VM, it does not matter whether the functionality is implemented in the hypervisor or in the VM. We can also confirm this observation by substituting $n = 1$ in inequality (10.3).

Figures 10.8(a) and (b) illustrate how R_F varies as R_f is increased from 0.1 to 1. The graphs show that for configuration C_2 to be more reliable than C_1, F has to be more reliable than f. Figure 10.8(a) shows that as $R_{M'}$ increases, the degree by which F should be more reliable than f also increases. Figure 10.8(b) shows that the degree is also considerably higher when more VMs are co-hosted on the same physical host. For example, even with modest $R_{M'}$ and R_f values of 0.75, F has to be ultra-reliable: R_F has to be more than 0.9932 and 0.9994 if $n = 6$ and $n = 9$, respectively. Thus, when more than a handful of VMs are co-hosted on the same physical node, a better system reliability is more likely to be obtained by retaining a poorly reliable functionality in the VM rather than by moving the functionality into the hypervisor.

Virtualized Node with VMM-level Voting: Consider a fault-tolerant 2-out-of-3 replication scheme in which three VMs providing identical service are co-hosted on a single physical node. The VMM layer receives client requests and forwards them to all three VMs in the same order. Assume that the service is a deterministic state machine; thus, the VM replicas yield the same result for the same request. The VMM receives the results from the VM replicas. Once the VMM has obtained replies from two replicas with identical result values for a given client request, it forwards the result value to the corresponding client. Such a scheme can tolerate the arbitrary failure of one VM replica, and is similar to the one suggested in the RESH architecture for fault-tolerant replication using virtualization [92]. Assuming that the VMs fail independently, the



(a) Fixed n, varying $R_{M'}$ (b) Fixed $R_{M'}$, varying n

Figure 10.8: Plot of $R_F \geq \dfrac{[1-(1-R_f R_{M'})^n]}{[1-(1-R_{M'})^n]}$

system reliability is given by

$$R_{sys}^{2-\text{of}-3} = R_H R_V [R_M^3 + \binom{3}{2} R_M^2 (1-R_M)].$$

Then, $R_{sys}^{2-\text{of}-3} > R_{sys}^{NV}$ gives the condition for the 2-out-of-3 replication scheme to be more reliable than the non-virtualized service. Thus, we obtain

$$R_H R_V [R_M^3 + \binom{3}{2} R_M^2 (1-R_M)] > R_H R_M$$

$$\implies R_V > \frac{1}{3R_M - 2R_M^2}. \tag{10.4}$$

Inequality (10.4) gives a lower bound on the hypervisor reliability for the 2-out-of-3 replication scheme to have better reliability than the non-virtualized case. Figure 10.9 shows a plot of $\frac{1}{3R_M - 2R_M^2} < R_V < 1$. It is clear from the graph that there exists no R_V value that satisfies inequality (10.4) and is less than 1 when $R_M \leq 0.5$. In other words, if the VM reliability (i.e., the operating system and service reliability) is poor to begin with, then the 2-out-of-3 replication scheme will only make the node reliability worse even if the hypervisor is ultra-reliable. This result concurs with the well-known fact that any form of redundancy with majority voting is not helpful for improving overall system reliability when the overall system is composed of modules with individual reliabilities of less than 0.5 [58]. The graph also shows that the higher the hypervisor reliability, the larger the range of VM reliability values for which the 2-out-of-3 replication scheme has better reliability than the non-virtualized case. For example, when $R_V = 0.98$, the range of VM reliability values that can be accommodated is greater than the range when $R_V = 0.9$.

10.5 An Architecture for a More Reliable Xen VMM

As shown by the model-based analysis in Section 10.4, it is highly desirable to make the VMM as reliable as possible to improve the overall reliability of a virtualized node. In this section, we leverage our X-Spy implementation to propose a reliability-enhanced design of the popular Xen open-source VMM [11].

The Xen VMM (Figure 10.10(a)) consists of a hypervisor core and a privileged domain (or VM) called Dom0 or domain zero. The hypervisor core is small in size and

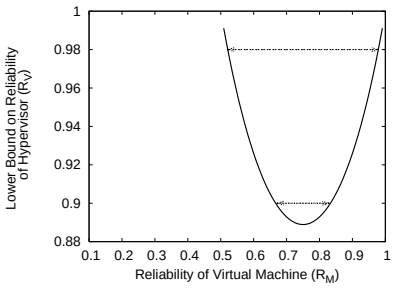
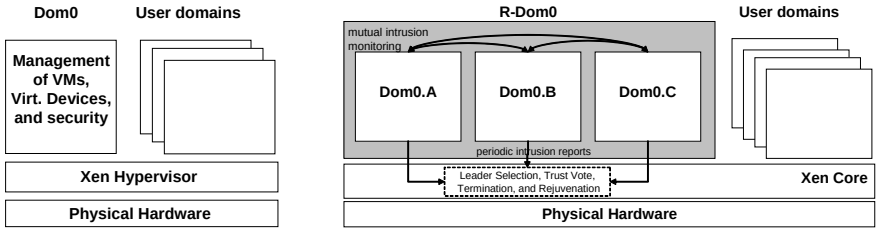


Figure 10.9: Plot of $(3R_M - 2R_M^2)^{-1} < R_V < 1$



(a) Normal Xen Architecture (b) R-Xen Architecture

Figure 10.10: Enhancing the Reliability of the Xen VMM

concerned with virtualizing the memory and CPU. Dom0 is a full-fledged VM running a guest OS (Linux) and virtualizes other hardware devices (such as disks and network interfaces). Dom0 is the first domain that is created, and controls all other domains, called user domains or DomUs. For any given physical device in Xen, the native device driver is part of at most one VM. If the device is to be shared with other VMs, then the VM with the native device driver makes the device available through a *back-end driver*. Any VM that wants to share the device exports a virtual device driver called the *front-end driver* to the back-end driver. Every front-end virtual device has to be connected to a corresponding back-end virtual device; only then does the front-end device become active. The mapping is one-to-one, i.e., each front-end virtual device from each user domain is mapped to a corresponding back-end virtual device. The communication between the back-end and front-end drivers takes places through shared memory and event channels. The event channel is used for sending simple lightweight notifications and the shared memory is used for sending requests and data.

As Dom0 is relatively large, we expect its reliability to be lower than that of the hypervisor core. Thus, improving the reliability of Dom0 is crucial to improving the reliability of the Xen VMM as a whole. We combine some of the technologies described in Section 10.2, namely, intrusion detection, enforcing fail-stop behavior, and intrusion response in form of software rejuvenation, to architect a more reliable Xen VMM, which we call *R-Xen*.

In R-Xen, we enhance the reliability of Dom0 by replication (Figure 10.10(b)). Dom0 is a single logical entity that actually consists of three privileged domains, *Dom0.A*, *Dom0.B*, and *Dom0.C*, with identical privilege levels. The three replicas mutually monitor each other using the techniques we described above in our X-Spy implementation. Specifically, each Dom0 replica is simultaneously the PVM and the

SSVM for the other two Dom0s. Periodically, the Dom0 replicas submit a fault detection vote to the hypervisor core that indicates whether one of its two peers is thought to be compromised. If any given Dom0 replica is labeled as being faulty by its two peer SSVMs, then the replica will be terminated and rejuvenated by the hypervisor. In this way, we enforce fail-stop behavior of the replica despite the presence of a more severe kind of fault in the replica. The hypervisor core then starts a new Dom0 replica as a replacement of the terminated one.

One of the Dom0 replicas is designated as *active* by the hypervisor core, and it is this active replica that provides the back-end drivers for the devices of the user domains. The other two replicas are designated as *passive*, and do not provide any back-end devices. As mentioned above, each of the three Dom0 replicas monitors and is being monitored by the other two. If the hypervisor gets reports from two independent replicas labeling the third replica as faulty, then the hypervisor terminates that replica and replaces it with a new Dom0 replica. If the terminated replica is a primary, then the hypervisor designates one of the backups as the new primary replica by re-connecting the front-end devices of the user domain(s) to the replica's back-end devices. The disconnection and reconnection of the user domain(s) to a different Dom0 has already been implemented in Xen and is used for live migration of domains. Therefore, the code can be reused. The hypervisor itself has to actively give permissions for doing the reconnection and re-routing the data from the old Dom0 to the new one. It also has to shutdown the old Dom0 after the reconnection process has been completed. Using a previously started backup as the new primary results in less interruption to the user domain than using the replacement replica (which has to be booted from scratch) as the new primary. It also enables the booting of the replacement replica to occur concurrently to the re-connection of the front-end devices. Like other fault-detection-based techniques, there is the drawback of *detection latency*, i.e., a time delay between the actual occurrence of the fault and its detection. I/O requests sent by the user domain(s) during this latency period may have to be re-issued. On the positive side, our technique can be implemented in a manner that is completely transparent to the user domain(s). In other words, a DomU running on normal Xen should be able to run without modification on this type of R-Xen as well.

Part III

Evaluation and Outlook

Chapter 11

Security Management Components of the OpenTC Virtual Datacenter Prototype

11.1 Overview

We realized a prototype of a physical data center implementing multi-tenant virtual data centers. Each one, usually owned by a different customer, is mapped onto a different TVD.

Our prototype is based on two classes of building blocks: Physical hosts providing VMs for the virtual domains and hosts running the data center services.

Figure 11.1 shows a host of the first class, a single secure hypervisor platform (that we henceforth refer to as "the platform") that is based on the Xen architecture [11] (Xen 3.1.1 and Linux kernel 2.6.22). Following the Xen model, the platform consists of a core hypervisor, the privileged management VM called Domain-0, and the guest VMs. We have also implemented the TVD Master role that runs on a second class host and maintains policies and membership for a given virtual domain.

The following components are provided by our platform. The *Compartment Manager* manages the life-cycle and the integrity of the guest VMs (also called compartments[1]). The *TVD Proxy* is the local representative of the TVD Master running on the platform and enforcing the policies of that TVD. Each TVD Proxy is created and spawned by a TVD Proxy Factory, whenever a TVD needs to be extended to a new host. The *Secure Virtual Network subsystem* creates, manages, and makes secure the virtual LAN for each TVD. The *Trusted Channel Proxy* implements the Trusted Channel needed for the pre-admission phase. In the current implementation, all these components run in Domain-0. Together with the Xen hypervisor and the other services in Domain-0, they constitute the Trusted Computing Base (TCB) of the platform, i.e., the integrity of these components must be ensured to guarantee the correct enforcement of the TVD policies.

The integrity measurements of the TCB components are performed during the au-

[1] A compartment is a protected execution environment, subject to information flow control policies enforced by the hypervisor: a compartment can securely host a single process running with threads as well as full-fledged VMs.

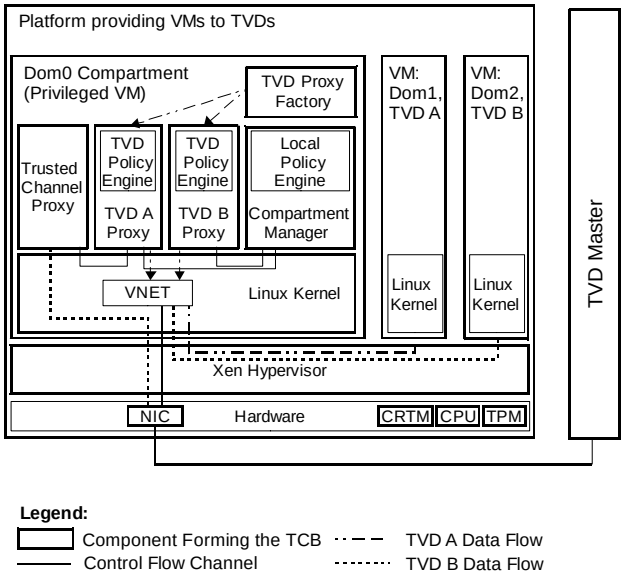


Figure 11.1: Xen Architecture for TVD.

thenticated boot, when the chain of trust begun by the Core Root of Trust for Measurement (CRTM) is extended using TrustedGRUB [4], a TCG-enhanced version of the standard GRUB boot loader. The TVD Master implements one end of the Trusted Channel establishment protocol. The other end is the TVD Proxy Factory.

Figure 11.2 shows the simplified layout of our prototype data center. It has two virtual networks per customer: a management network for managing the customer's VMs within the TVD and a user data network for the actual communication between client VMs. Furthermore, the data center has a DMZ for external connections, a virtual data center management network (VDC) that is used for communication between our data center management components, and finally a SAN network that provides storage to the individual platforms.

Each virtual domain has a intra-domain management VM that runs the management software of the customer and connects to the management network. This management software interacts with a filtered XenAPI (called XenAPI') that hides infrastructure and machines belonging to other virtual domains and provides a virtual view of the data center. Each administrator in a virtual domain can then define and start any number of guest VMs that get connected to the corresponding data network.

11.2 Security Services

In [55], we have described a Xen-based prototype of our security services for integrity management and obtaining compliance proofs. The prototype enables the protection of security policies against modification and allows stakeholders to verify the policies actually implemented. The paper also describes multiple use cases, in which we demonstrated the policy enforcement and compliance-checking capabilities of our implementation. For example, we showed how to validate the configuration of the virtual networking subsystem on each host (assuming that a TPM is embedded in each host).

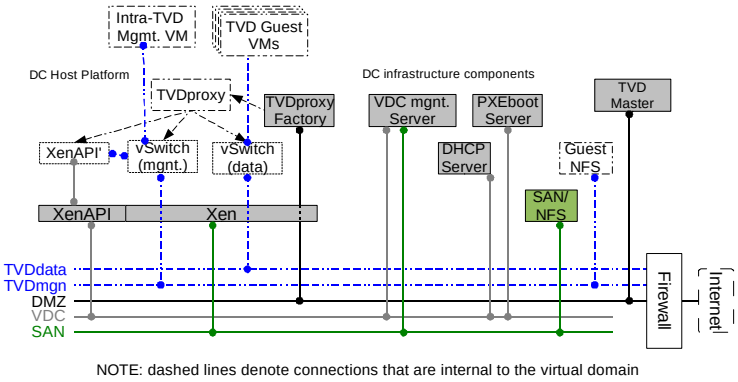


Figure 11.2: Layout of our prototype virtual data center.

Here, we provide an overview of our security services implementation. The Compartment Manager (CM) is responsible for the VM life-cycle management. As the sole entry point for VM-related user commands, it communicates directly with the hypervisor and orchestrates the Integrity Manager and the secure device virtualization functions of the platform. These functions base on the standard Xen management tools.

The CM through the Integrity Manager (IM) is responsible for verifying the integrity of the VMs. The root file system is made available to each VM, including Domain-0, through a pair of partitions or virtual disks. One of them is read-only, contains the initial root file system, and is measured together with the VM configuration file; the resulting whole integrity measurement of the VM is then accumulated into the TPM by extending one PCR. The other partition or virtual disk, which is read/write and empty at the beginning of the VM life, records the changes occurring on the root file system while the VM is running. This is done through the copy-on-write mechanism provided by *unionfs* [5, 1] which allows the stacking of multiple file systems and provides the operating system with a unified view. CM and IM are also responsible for guaranteeing the confidentiality of a VM's data (stored on the second read/write image) by encrypting it using *dm-crypt* [2], which is the Linux device mapper with support for encrypting block devices (e.g. physical or virtual disks and partitions). Next, the encryption key is sealed against the measurement of the TCB and of the VM (stored in a set of PCRs). We use this "sealed disk" scheme to protect the VM's confidential data. This scheme can be applied to all disks, except the root image, depending on the security requirements given.

The TVD policy defines, for each VM, the disks that need to be measured, the PCR(s) in which the measurements need to be stored, and the disks that need to be sealed. Once the policy has been executed and all disks have been prepared (measured/unsealed), the admission protocol involving the CM and the TVD Proxy (see Sections 3.4 and 11.3) follows. Then the CM requests Xen to start the VM. To manage VMs, the CM maintains the association between a running VM and the policy that was used to start it.

11.3 TVD Master, Proxies and Proxy Factories

The TVD policy of our prototype lists all VMs that can potentially be admitted to the TVD. Each VM is identified by the Xen domain identifier. For each VM, the policy specifies the required integrity measurement value. Only if the VM's actual measurement value (obtained by the CM) matches the required value the VM will be admitted to the TVD. The policy also specifies the MAC address assigned to the VM's virtual NIC, if the admission is successful. Moreover, it identifies the VLAN corresponding to the TVD, and the VPN keys needed for establishing secure connections within the TVD. Currently one shared key per TVD is used to protect all insecure links.

When a VM wants to be admitted to the TVD (i.e., this is stated in the configuration file of the VM), the related TVD Proxy is contacted by CM using a stateless TCP-based protocol called Compartment Admission Protocol (CAP). Since a platform can host VMs belonging to different TVDs, the CM contacts the TVD Proxy Factory to obtain the TVD Proxy end-point for the requested TVD. If such TVD Proxy is not running yet, the TVD Proxy Factory creates and spawns it. Before starting the VM, CM measures it (as explained in Section 11.2) and asks the TVD Proxy whether the VM can be admitted to the TVD by passing the identifier and the measurement of the VM. If the answer is positive, CM receives the MAC address specified in the policy from the TVD Proxy, creates the back-end network device (see Section 11.4 for further explanations about Xen back-end and front-end devices), and sets the MAC address for the front-end device. Finally, the CM requests the TVD Proxy to attach the VM to the virtual switch (vSwitch) of the VLAN corresponding to the TVD by specifying the identifier, measurement value, and the names of back-end devices for the VM being admitted. In the case of a negative response from the TVD Proxy, the CM can be configured to either start the VM even though it will not be connected to the TVD VLAN or not to start it at all.

11.4 Secure Virtual Network subsystem

The prototype implementation of our Secure Virtual Network subsystem has been documented in [18] and has been integrated in the prototype being presented in this paper. Our networking extensions consist of vSwitches, VLAN tagging, and LAN encapsulation modules. They are implemented as kernel modules in Domain-0, which also acts as the driver VM for the physical NIC(s) of each physical host.

To specify the particular vSwitch and the particular port in the vSwitch to which a VM's Xen back-end device must be attached, the Xen VM configuration file is used. This file is generated by CM after having received information (MAC address and VLAN identifier) from the TVD Proxy. We use additional scripts to specify whether a particular vSwitch should use one or both of the VLAN tagging and encapsulation mechanisms for isolating separate virtual networks.

The vSwitches maintain a table mapping virtual network devices to ports on a particular vSwitch. The encapsulation module implements EtherIP processing for packets coming out of and destined for the VMs. The VLAN segments associated with different TVDs and the corresponding vSwitches are assigned unique identifiers. The network identifier field in the EtherIP packets is set to the identifier of the vSwitch that the target VM is attached to.

The VLAN tagging module tags the packet with the VLAN identifier corresponding to the VLAN that the target VM is a part of. At the destination platform, the VLAN

module removes the VLAN tags, and routes the packets to the appropriate vSwitch based on the VLAN tag.

11.5 Trusted Channel Proxies

The Trusted Channel between TVD Proxy Factory and TVD Master used during the pre-admission phase is set up by means of a pair of services called Trusted Channel proxies. They implement the Trusted Channel at the application layer via a TLS tunnel, made available to TVD Proxy Factory and Master once remote attestation has been successful. The remote attestation is done by performing the `TPM_Quote` operation, namely, digitally signing a set of PCRs and a challenge received from the remote attester using a TPM asymmetric key. The latter can be certified as Attestation Identity Key (AIK) by a Privacy CA. The result of the `TPM_Quote` operation (i.e. the signature over a set of PCR values), the actual PCR values and the AIK Public Key Certificate are sent by the TVD Proxy Factory to the TVD Master to be verified. If the verification is successful, then the remote attestation can be considered successful, and the two proxies start tunneling the incoming TCP packets through the TLS channel. An alternative implementation for attesting via Trusted Channel is documented in [9] and will be integrated in our prototype. This approach replaces the `TPM_Quote` operation with the usage of sealing and TPM certified keys.

Chapter 12

Evaluation of the Prototype

12.1 Performance Evaluation

We implemented our data center prototype using HP ProLiant BL25p G2 blade servers each fitted with two AMD Opteron processors running at 2 GHz, 8GB system memory and a Gigabit Ethernet card. We now discuss the performance of our prototype. Most of our components implement or support management tasks. They are dedicated to automate the secure set-up of platforms, virtual machines and domains, if possible with minimal performance impact on the running system.

Note that measuring performance of virtual systems is not straightforward. We first used `iostat` to retrieve CPU usage data from `/proc/stat` under Linux and in Domain-0. This wrongly indicated that Xen needs half the CPU as compared to Linux. We then used the Xen tool `xentop` that gathers the data via hyper-calls to the hypervisor. Since this represents the time given to each VM (Domain-0, guest VM) by the hypervisor, the resulting data was no longer distorted by the virtual notion of time given to the VMs in Xen.

12.1.1 Management

In Table 12.1 compares the boot-up and shut-down delay between for virtual machines using the unmodified Xen implementation and our components. Averaged over 235 measurements, our components add some 10 percent to the the original time-to boot. The delay is caused by measuring the configuration, attesting to the VM, transferring and checking configuration measurements, and (in case of success) attaching the VM to the corresponding network. Stopping a virtual machine requires 2.4s instead of 0.5s for the original Xen implementation. Here, the overhead is caused by the fact that the current implementation of the compartment manager polls the VM in rather long intervals to verify a successful shut-down.

		System Measured	
Operation		**Prototype**	**Xen**
Management	Start	3.601s	3.332s
	Stop	2.371s	0.460s

Table 12.1: Performance Measurements of our Prototype

Throughput	Linux	VLAN Tagging	Xen Bridging	EtherIP
TX (Mbps)	932	883	872	847
RX (Mbps)	932	881	870	851
Comparison	100%	56.79%	55.98%	54.69%

Table 12.2: NetPerf Benchmark: Guest VM to Guest VM Throughput.

	Minimum	Average	Maximum	Mean Deviation
Bridged	0.136s	0.180s	0.294s	0.024s
VLAN	0.140s	0.212s	0.357s	0.030s
EtherIP	0.151s	0.246s	0.378s	0.034s

Table 12.3: Round-trip Times using Ping.

12.1.2 Networking

We obtained the throughput results using the `NetPerf` network benchmark and the latency results using the `ping` tool. Using the former benchmark, we measured the Tx (outgoing) and Rx (incoming) throughput for traffic from one guest VM to another guest VM on the same physical host. To do so, we ran one instance of the benchmark on one guest VM as a server process and another instance on the second guest VM to do the actual benchmark.

We report the throughput results for different networking schemes in Table 12.2. The figures show that the throughput results for both VLAN tagging and EtherIP schemes are comparable to that of the standard Xen (bridge) configuration. As expected, VLAN tagging yields the best throughput in a virtualized system that outperforms the standard Xen configuration. Both Xen bridging and VLAN tagging perform better on the **Tx path**. For EtherIP, the major cost in the Tx path is having to allocate a fresh socket buffer (`skb`) and copy the original buffer data into the fresh `skb`. When first allocating a `skb`, the Linux network stack allocates a fixed amount of headroom for the expected headers that will be added to the packet as it goes down the stack. Unfortunately, not enough space is allocated upfront to allow us to fit in the EtherIP header; so, we have to copy the data around, which is very costly.

In the Rx path, there is no packet-copying overhead for the EtherIP approach; the extra EtherIP header merely has to be removed before the packet is sent to a VM. As compared to VLAN tagging, in which packets are grabbed from the Linux network stack, EtherIP requires that packets are passed to and processed by the host OS IP stack before they are handed over to the EtherIP packet handler of the vSwitch code.

Table 12.3 shows the round-trip times between two guest VMs on a physical host for the bridged, VLAN, and EtherIP encapsulation cases obtained using the `ping -c 1000 host` command, i.e., 1000 packets sent. The results show that the average round-trip times for VLAN and EtherIP encapsulation are 17.8% and 36.7% higher than that of the standard Xen bridged configuration.

12.1.3 Storage

Timing the set-up of storage has been part of the management evaluation. We now evaluate actual run-time performance of virtual disks.

We compared three set-ups using an IBM Thinkpad T60p: Disk access from Dom0,

CPU Utilization	Linux	Xen (Dom0+DomU)
Encrypted	42%	45% (42+3%)
Unencrypted	5%	13% (10+3%)

Table 12.4: CPU Utilization of Virtual Disks at 30MB/s.

disk access in a Linux without Xen, and disk access from a guest VM. For each of these three systems we compared encrypted and unencrypted access.

We first measured read/write throughput. A first observation was that in all cases, the disk performance limited our overall performance, i.e., encryption did not result in a performance penalty in a single-VM system (all 6 set-ups provided approx 30MB/s throughput).

As a consequence, we then measured the CPU utilization of the different set-ups (see Table 12.4). This figure points out that encrypting a disk at 30MB/s requires 42% CPU under Linux and 45% under Xen while servicing a guest VM. This shows that the utilization is similar to a plain Linux. The fairly high CPU utilization substantially limits the usage of encryption in a data centers. Fortunately, encryption in a data center can often be replaced by physical security or other measures. The only exceptions are removable media that are moved between locations.

12.2 Limitations of our Prototype

OpenTC has made substantial progress in design an implementation of secure virtualization. However, there are still open challenges to overcome.

Standardized Management APIs One goal of our prototype is to enable independent and transparent[1] management of the different virtual domains. In principle, we aimed at allowing each TVD administrator to manage a given TVD independently of all other domains. This would require that a TVD owner can define policies, define images, and start them while being independent of others and not being required to modify any tooling. While our prototype achieves management independence, we require some modifications to existing tooling. While machines can be managed using the standard APIs libVirt or XenAPI (see `libvirt.org` and `wiki.xensource.com/xenwiki/XenApi`), there exists no similar API for managing virtual machine images. As a consequence, we require that each customer uploads images into per-domain sub-directories. Once the images are stored, they can then be referenced during machine creation.

Federated Identities The first limitation are globally unique identifiers. Our current prototype uses a naming authority. In the long run, TVDs should be federated without any mediation of a central authority, making an identification scheme like UUID [72] necessary to guarantee the uniqueness beforehand.

Integrity Management Architecture The current scheme for measuring VM integrity is coarse-grained, because the entire file system is measured.[2] It is a first at-

[1] With transparent, we mean that each customer can use unmodified standard APIs to manage its machines and resources.

[2] Note that unlike [103] we aimed at implementing integrity-protection at the virtualization layer.

tempt of measuring the VMs while allowing a persistent storage; however it has a big shortcoming: the measurements do not capture the modifications that occurred on the file system because they are stored on the second read/write virtual disk, which is never measured. Moreover, the VM is measured only prior to being started, and so far there is no support for run-time monitoring yet. In the long run, we will use a finer-grained integrity measurements, e.g., through a virtualization-enhanced version of the IMA proposed in [103] while using integrity-enhanced storage [24, 89] to protect data at rest. However, these concepts are not widely available and come with a performance penalty. In the long run, integrity-protecting storage hardware would resolve these issues[3].

Policy Management and Distribution Another part of our architecture that has not been fully implemented is the TVD Masters. Today, they only perform intra-TVD policy distribution. In the long run, they should enable trust brokering and delegation to allow trust establishment between multiple TVDs.

Finally, in this first implementation, all TVD components reside in Domain-0, the Xen-privileged VM. Following the approach of Domain-0 disaggregation[4] [83], the TVD Proxy and VNET components will be moved away from Domain-0 to run in dedicated and isolated VMs.

Secure Virtual Networking Our networking approach adds security mechanisms in Dom0 of Xen. In the long run, virtualization support in the networking hardware would be desirable to reduce context switches and thus increase performance. Note that this requirement is independent from security. However, for securing network, VPN support in the network cards would be desirable in particular for datacenters.

[3]With integrity-protecting storage we mean that the hard disk can output a hash-value for a range of blocks. Integrity-protection then means that only blocks corresponding to this hash-value are readable while modified blocks will produce read errors.

[4]The OpenTC consortium is pursuing this approach to reduce the weight of the trust assumptions on Domain-0.

Chapter 13

Conclusion and Outlook

13.1 Lessons Learned

Embedding integrity verification mechanisms in a distributed security architecture creates serious challenges. Conceptually, many of them can be addressed with property-based attestation. However, property-based attestation depends on well-defined and potentially institutionalized processes for validating the behavioral equivalence of different configurations. Certifying that two configurations have identical properties is currently a manual and labor-intensive exercise, which is costly and does not scale beyond single TVD owners or data center administrators.

While the migration of VMs between different physical hosts is well understood, the migration of a complete trust context associated with a VM has proved to be difficult. The latter type of migration requires the migration of not only the virtual networking and storage devices (with associated cryptographic keys, if necessary), but also of a virtual TPM, if present, which will be rooted in different hardware TPMs prior to and after the migration. During the migration process, the integrity of all these components has to be guaranteed while managing the handing-off of device access in a transactional fashion. Note that the importance of securing this transition has been further emphasized by recently published attacks on virtual machines in transit.

Building a security API that is at the same time flexible, usable, and manageable has proved to be more difficult than expected. A key reason for this difficulty is the requirement that the API should be easily adaptable to other hypervisor architectures and to workstation scenarios with GUI interfaces. While addressing each of these requirements separately is feasible, their combination comes with many trade-offs.

Yet another type of trade-off concerns our aim of decomposing the Xen architecture into multiple security services each running in dedicated tiny VMs while reducing the reliance on Domain-0, the privileged management VM. While such decomposition is advantageous from a security perspective, it tends to reduce flexibility. The availability of a full-fledged Linux management VM with access to all subsystems enables easy extensibility and rapid prototyping (scripting, adding devices, firewalls, VPNs etc), and also corresponds to the expectations of many Xen users. In general, however, considerations of usability tend to favor design decisions that are sub-optimal from a strict security perspective.

13.2 The Future of Secure Virtualization and Trusted Computing

WP5 aimed at securing virtualization by using trusted computing technologies. We managed to prototype a secure virtual datacenter that uses trusted computing hardware for proving its integrity to stakeholders.

13.2.1 Trusted Computing

Observations During the course of our project, we have made several observations. The first is that trusted computing hardware (i.e., TPM) is useful for protecting keys and cryptographic operations using these keys. This usage is establishes a hardware-protected machine identity. This is similar to today's usage of smart-cards.

We furthermore realized that integrity-protection using this hardware is very difficult in practice. Due to the fact that minimal changes to the software render an attestation invalid, maintaining the integrity of such a system is error-prone.

We believe that there are several aspects that make integrity-validation challenging. The first is the variety and variability of code. Even in our security project the code-based changed a lot. As a consequence, any practical integrity-validation mechanism needs to be able to tackle this code evolution while distinguishing between desired (=secure) and undesired (=insecure) upgrades. One approach towards solving this problem is to use the Linux package management infrastructure to define acceptable code [82].

The other challenging aspect is the lack of isolation in today's environments. Today, any malicious code in a VM can essentially take over a complete VM. As a consequence, the integrity of any code in the VM depends on all other code. As a consequence, integrity of a component can usually not be evaluated in isolation. This substantially increases the overall complexity of an integrity assessment since in principle any untrusted component on the system threatens overall integrity of most other components. It also lowers the performance since any change in a VM needs to be assessed. While we provide isolation between VMs, a long-term solution would be to lower the granularity to provide strong isolation on an object or package level.

Consequences We believe that the pervasive deployment of TPM hardware will further enable and support the usage of TPMs as machine identities and for protecting cryptographic keys. However, based on our observations, we believe that the applicability of trusted computing in the near future will be limited to small-scale and stable high-assurance systems. Examples include embedded systems, or high-assurance systems such as high-assurance workstations for defense or very integrity-critical applications. Another example are virtualization platforms for executing virtual appliances. For these cases, the limited and stable trusted computing base is effectively verifiable using trusted computing.

Open Challenges The first open challenge is to provide and manage finer-grained isolation to enable more localized integrity verification of individual components. This would enable tree-like verification of a trusted computing base where siblings cannot destroy each others integrity. This is similar to the OpenTC approach of running independent virtual machines that cannot invalidate each others integrity. However, for an actual operating platform based on these principles, the granularity would need to

be lowered from the VM to the object level. Overall, this concept is closer to the isolated-objects approach of L4. However, for practical applicability the performance and life-cycle management of thousands of isolated computing objects will remain challenging.

A second open challenge is run-time attestation, i.e., moving from boot-time validation of a system towards real-time validation. This means that whenever code is executed that this code has been validated. Without such near-time validation, in particular long-running systems that were unmodified at boot-time can be attacked at run-time such that the code changes are not detected.

A third open challenge is to develop new approaches towards integrity validation that are based on the integrity pedigree of isolated objects. We believe that such a method will require a high amount of initial validation but should then enable simplified updates to maintain running code along with its integrity assurance.

13.2.2 Secure Virtualization

Observations In OpenTC, we have implemented a virtual datacenter that shows that isolated machines with well-defined sharing. We managed to automated most management tasks. This enabled us to hide implementation details from users such that users could concentrate on their computing resources that were independent of the resources of other customers. Automating management of security has been essential since manual management is costly and error-prone.

Consequences We believe that there is a trend towards such scalable virtual infrastructures that are automatically managed and hide its implementation details from users. This so-called “Cloud Computing” approach has clear advantages from a usability perspective. Unfortunately, today’s cloud computing platforms do not provide sufficient security. Furthermore, from a security perspective, hiding implementation details no longer allows users to validate the security of the platform themselves.

Overall, we believe that this in-transparent approach towards cloud computing will be limited to low-trust scenarios similar to today’s grid computing. These clouds will either be used within organizations that trust each other (within an enterprise or with a strong supplier-customer trust relationship). Another scenario are uncritical applications such as web-servers or scientific computing on public data.

Open Challenges An open challenge is how to balance transparency with security assurance. In OpenTC, we aimed at full verifiability, i.e., a customer can see all implementation details of the infrastructure. Only actual configurations of peer customers are hidden from our users. In the long run, however, hiding these implementation details (which are secrets of the provider) from customers while guaranteeing security will remain challenging. In addition, secure composition of virtual services that are potentially provided by different cloud implementations operated by mutually mistrusting providers will remain a challenge. In particular for this cloud-of-clouds scenarios well-defined operation and implementation guidelines (and audits) as well as standardized assurance exchange interfaces will be required [61].

13.3 Conclusions

In this report, we have described the OpenTC approach and components for a Trusted Virtual Datacenter. The core idea is to build a datacenter that allows for and isolates multiple customers while satisfying the security requirements of each customer.

Securing the access to data on persistent media and during transfer over the network is a serious problem in distributed virtual data-center and cloud computing scenarios. We described a framework based on TVDs and Trusted Computing for secure network and storage virtualization that includes mechanisms for verifying the integrity of the hypervisor, VMs, and security policy enforcement points.

This is achieved by means of so-called Trusted Virtual Domains that provide a virtual environment for each customer. Each trusted virtual domain exposes a standardized management interface that allows customers to run existing management software to manage their respective domains. The concept of TVDs is rigid enough to allow consistent policy enforcement across a group of domain elements, while being flexible enough to support policy-controlled interactions between different TVDs. TVD policies and configurations are 'backward-compatible' in supporting options that could be taken for granted in non-virtualized data centers. For example, co-hosting of specific customer services with those of other data-center customers on the same physical platform could be inhibited if so desired. By incorporating hardware-based Trusted Computing technology, our framework allows the creation of policy domains with attestable trust properties for each of the domain nodes.

In order to enforce the given security requirements, we deploy the security services described in Deliverable D05.4 [88] that allow verifiable security within each domain. This includes validation of the trusted computing base as well as verifiable enforcement of given per-domain policies. The inclusion of integrity measurement and management mechanisms as part of the physical platform's TCBs provides both data-center customers and administrators with a much needed view of the elements (hypervisors, VMs, etc.) that are part of their virtual infrastructure as well as information on the configurations of those elements. Our framework can be used to obtain information about the elements on a 'need-to-know' basis without having to introduce all-powerful roles of administrators with access to every aspect of a platform.

We have substantially contributed to the knowledge in secure virtualization and trusted computing. Overall, we managed to build the first large-scale virtual datacenter that can be validated using trusted computing technology. Nevertheless, despite our substantial progress, there are still several open research challenges. Examples include large-scale management and validation of integrity as well as efficient implementation strong isolation guarantees on a finer-grained level. In the long run, we hope that this will enable to lower the granularity of integrity guarantees from Virtual Machines to individual objects.

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