TOWARDS MORE SECURE PROGRAM EXECUTION ENVIRONMENTS

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Contents

Summary

The increasing prevalence of cyber attacks is a worrying trend in the Internet age. By exploiting vulnerabilities in operating systems or applications, intruders are quite often able to circumvent the existing security mechanisms. This thesis proposes measures and infrastructure to provide more secure program execution environments so as to enhance host security. Our approach is based on securing the "Program Protection Life Cycle (PPLC)", which protects application programs throughout their life cycles against attacks, including zero-day attacks. A number of security mechanisms are proposed along the PPLC to substantially reduce attack vectors on a host.

Firstly, to mitigate the threat of zero-day attacks to a running program, we investigate a system-call monitoring Intrusion Detection System (IDS) which aims to detect any potential anomalous behavior of the execution. Using an automated attack generation approach, we show how a non-parameterized Self-based IDS model is vulnerable to mimicry attacks. We then propose an improved IDS model to mitigate mimicry attacks by employing a privilege and argument abstraction technique. We also move on to propose a general framework based on a notion of "attack-space search" to demonstrate how the attack construction approach can apply to various IDS models. The framework is then used to measure the resistance level of IDSs against attacks targeted on them.

Secondly, to secure program invocations on a host, we propose a lightweight executable authentication scheme which provides secure program distribution and integrity assurance on the invoked program. This is complemented by an automated vulnerability management scheme, which is aimed at performing automated vulnerability checks on operating system components and application programs to ensure vulnerability-free executions.

Thirdly, we address a supporting infrastructure which is needed to provide an efficient and secure program distribution and associated usage. Since existing Public Key Infrastructure (PKI) certificate revocation mechanisms are not sufficiently lightweight and timely, we propose a family of lightweight and practical near real-time revocation schemes. Our schemes are based on the use of the recently available Extended-Validation Certificate infrastructure, and can offer timeliness guarantees on the order of minute(s) with low performance cost. We also propose a formalism to reason with PKI-based systems and protocols by enhancing BAN Logic to deal with modern PKI-based protocols.

In summary, the contribution of this thesis is to give additional layers of protection, which give greater assurances of secure program executions amidst the increasing malware threats in today's Internet-connected systems.

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List of Abbreviations

CA	Certificate Authority
CMAE	Certificate Management Ancillary Entity
COTS	Commercial Off The Shelf
CPE	Common Platform Enumeration
CRL	Certificate Revocation List
CRS	Certificate Revocation Scheme
CRT	Certificate Revocation Tree
CSI	Certificate Status Information
CVE	Common Vulnerabilities and Exposures
DoS	Denial of Service
EV	Extended Validation
EVC	Extended Validation Certificate
FSA	Finite State Automaton
IETF	Internet Engineering Task Force
IDS	Intrusion Detection System
IPS	Intrusion Prevention System
LFC	Locality Frame Count
MAC	Message Authentication Code
MHT	Merkle Hash Tree
NVD	National Vulnerability Database
OCSP	Online Certificate Status Protocol

OS	Operating System
OSVDB	Open Source Vulnerability Database
OVAL	Open Vulnerability and Assessment Language
PAC	Privilege and Argument Categorization
PC	Program Counter
PDA	Push Down Automaton
PDF	Probability Density Function
PKI	Public-Key Infrastructure
PPLC	Program Protection Life Cycle
SCAP	Security Content Automation Protocol
SSL	Secure Sockets Layer
TLS	Transport Layer Security
TCG	Trusted Computing Group
TPM	Trusted Platform Module
UAC	User Account Control

Chapter 1

Introduction

Malicious software (malware), which is designed to infiltrate a system and cause damage to it, is a critical threat today. A report by F-Secure [42] indicates that there was as much malware produced in 2007 as in the previous 20 years altogether. A report by OECD [120] additionally highlights a worrying trend that malware has now evolved from occasional exploits to a global multi-million dollar criminal industry, and is threatening the Internet economy. On the other hand, computer systems, particularly software, are much more susceptible to attacks now than in the past. Hoglund and McGraw [61] attribute three factors to present-day software security problem, namely: increasing connectivity, complexity and extensibility.

Despite numerous built-in security measures deployed by modern operating systems (OS), the growing number of successful attacks shows that quite often intruders are able to circumvent the measures by exploiting flaws in the OS or applications. To mitigate this situation, additional measures which can be deployed on top of the existing OS security mechanisms are needed in order to harden the OS and provide an increased system protection. In addition, related infrastructure, such as Public Key Infrastructure (PKI), which critically supports host security mechanisms must be reliably available.

1.1 Securing Program Execution Environments

The central theme of this thesis is how to enhance host security by providing more secure environments for program execution. This thesis focuses on a concept of protecting software, i.e. programs, to ensure that they run as intended without violating host security. In particular, we focus on the challenges posed by commercial-off-the-shelf (COTS) software where no source code is available. Our aim is to safeguard software, from their distribution to their execution, from a variety of attack by the attackers.

We secure a program's execution environments based on our model of "*Program Protection Life Cycle (PPLC)*". The final objective of PPLC is to establish a belief on

host H concerning *good_intended_execution* property of a program P. That is, a host H believes that P performs the operations as intended by its (trusted) developer D, and that the operations will not violate H's security policies.[1] Figure 1.1 shows the overall components and life cycle of the PPLC model.

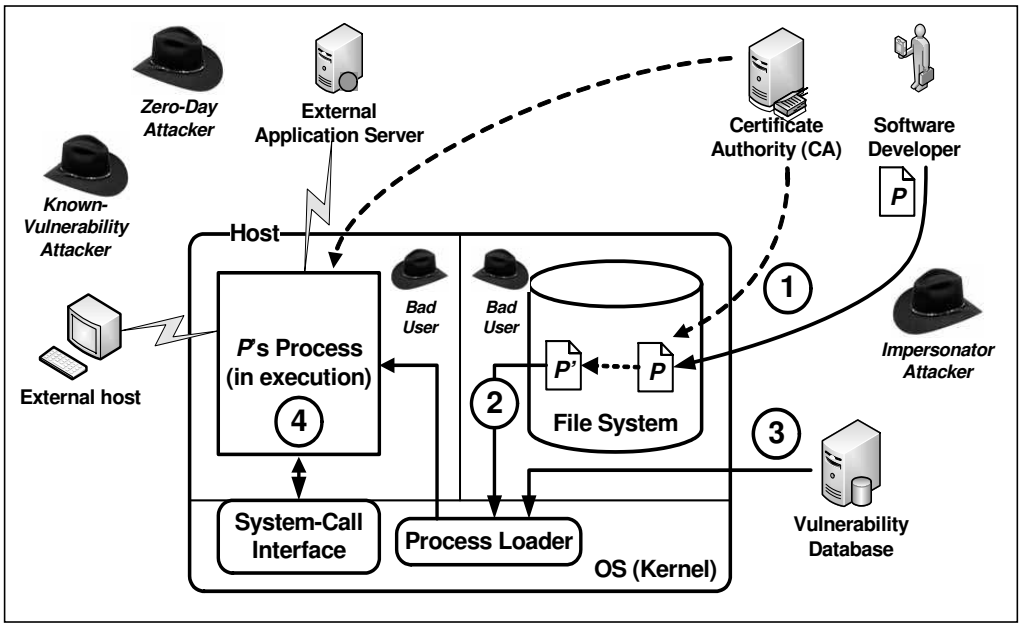


Figure 1.1: *Program Protection Life Cycle*: securing a program and its execution.

The four main steps of PPLC are:

1. **Secure program distribution**. This step is to give assurance to host H that program P originates from software developer D, and it is received unmodified. This goal can be achieved using public-key based signing on the program. Given a strong cryptosystem being used, the security thus relies on ensuring that the public and private key pair of D is valid (unrevoked) when H verifies P. By installing P into its system at time $t_{installed}$, host H implicitly agrees that P is a *good program*[2] which if it is run, it will not violate H's security policies. Hence, besides its content, P actually also carries with itself a belief statement that it is a good behavior program.

2. **Preserved program integrity for execution**. To protect system security, it is imperative to ensure that both the *content* and *pathname* of P at time $t_{invoked}$, just prior to its execution by the OS, is the same as those in $t_{installed}$. In other words, P's location and content must stay intact on the file system of H.

[1] Chapter 8 discusses logic systems, such as Burrows-Abadi-Needham (BAN) Logic [25, 26], which deal with beliefs pertaining to security properties on network protocols and computer systems.

[2] More specifically, by "*good program*", we mean that the program is non-malicious in nature, and has no intention to violate any security policies of the target hosts (beyond the program's known functionalities) or any acceptable use policies.

3. **Vulnerability-free program execution**. Despite the deployment of numerous security measures, a host becomes highly susceptible to attacks when executing P with publicly known vulnerabilities during its "*unpatched time interval*". This interval spans from the time when a vulnerability alert affecting P is published until patches are applied or other preventive/corrective measures taken by users. Given the increasingly high volume of reported vulnerabilities in recent years, the advantage is clearly with the attackers. To ensure that P runs as intended without any compromise due to attacks exploiting known vulnerabilities, automated vulnerability checking on P prior to program execution is important.

4. **Intrusion-free (unaltered) process execution**.

   Although P is believed to be good for execution, it still needs to be protected from possible attacks exploiting previously unknown vulnerabilities, which are also known as zero-day attacks. One useful mechanism to deal with this is Anomaly Detector Intrusion Detection System (IDS), which has an ability to detect novel attacks by flagging any abnormalities in program behavior. Online IDS systems which are well-integrated into the OS kernel can also block any suspected operations from being executed, thus preventing any potential security breach on H.

Safeguarding the PPLC steps brings two main benefits for host security. Firstly, it reduces or even possibly eliminates the need for some defense mechanisms, such as: anti-virus scanning on the executables, or network-packet scanning for known attack patterns. Secondly, it naturally deals with multi-modal attack entries. Program integrity protection, for instance, prevents file modification or addition either due to social engineering techniques, illegal user's operations or various malware activities. Vulnerability checking also prevents vulnerable programs from being executed, therefore inhibiting various modes of attacks exploiting known vulnerabilities.

Based on the PPLC model, in this thesis we thus argue the following statement:

> *"Security measures to protect program life cycle are important to establish, and that they will reduce many attack vectors on a host and ensure more secure program executions on that host. With the right approach and techniques, these measures can be deployed with acceptable performance cost while substantially increasing host security."*

1.2 Challenges in Securing Program Protection Life Cycle

There are a number of challenges faced in securing PPLC. We mention several main challenges, which have provided motivations to our work reported in this thesis.

1.2.1 Difficulty in Evaluating the Security of IDS

In IDS research, it is common for a proposed IDS to be shown, usually using several known attacks and some dataset, for being able to detect the attacks with high detection rate while exhibiting low false positive rate. The question of *how secure* the IDS against attacks on itself is usually overlooked. A "smart attacker", who know all the IDS internals, can exploit IDS limitations to craft attacks targeted on the IDS. One example highlighting this issue is the possibility of mimicry attacks [167, 152] on system-call monitoring IDS called Stide [60, 140, 141]. Hence, one step towards IDS security analysis is finding a systematic way to measure the *resistance level* of an IDS against attacks targeted on itself. The derived *attack generation time* is particularly useful to estimate how fast an attacker can come up with a stealthy exploit given a newly discovered vulnerability on the protected program.

1.2.2 Making Anomaly Detector IDS More Secure

Given possible attacks on anomaly detection IDSs such as Stide IDS [60, 140, 141], the question to be asked is how we can improve the IDS, as well as other IDSs sharing similar basic model (e.g. [136, 45]), to prevent the attacks. One approach is to make use of all available information from the observable events, such as system call arguments and user credential information. While increasing the IDS robustness against attacks, the inclusion of these new details should not prohibit fast (preferably online) detection operations, or increase the false positive rate.

1.2.3 Practicality of OS-based Executable Integrity Protection

The concept of program authentication, i.e. checking the integrity of stored executable programs, is not new. Tripwire [74] is one of widely known systems that provide (offline) file integrity protection. In addition, there are also a number of OS-integrated executable authentication implementations, although mostly on Unix [8, 170, 162]. Despite the potential great benefits brought, current commercial OSes, such as Windows and Solaris, do not generally come with built-in support for program integrity yet. One challenge here is to investigate how to devise a mandatory in-kernel executable integrity checking scheme, conducted just prior to program execution, in a lightweight manner on a real-world complex OS like Windows.

1.2.4 The Need to Automate Vulnerability Alert Processing

In order to keep track of security alerts on relevant OS and applications, system administrators usually rely on various sources of narrative vulnerability alert repositories. Given the speed at which an exploit becomes available once a vulnerability is known, and the

frequency of occurrence of such alerts, manual intervention is too slow, time-consuming and possibly ineffective. It is therefore a real challenge on how to develop a coordinated vulnerability database which is designed to be machine readable and processable. This will allow automated processing of alerts and corresponding vulnerability scanning to be done on a host. One particular issue is that the scanning process should provide assurance that it does not perform any operation beyond what is necessary. Thus, the scanner operations must be minimal and open to inspection.

1.2.5 Providing a Lightweight and Near Real-Time Certificate Revocation Scheme

Many host security mechanisms, including validating a digitally-signed program, rely on ensuring that the signer's public key is not revoked. In X.509-based PKI, certificate revocation is mainly supported by a mechanism called Certificate Revocation List (CRL) [36]. However, CRL is widely perceived to be costly [88, 58]. Several alternatives have been proposed, such as Online Certificate Status Protocol (OCSP) [113], which requires the Certificate Authority (CA) to respond to any incoming query in real time with a signed message. To secure software distribution and PKI-based program's interaction with external hosts, certificate revocation service must provide a *near real-time* timeliness guarantee and enable fast computation by a verifier. This is particularly important given the proliferation of lightweight computers and mobile devices. In addition, the revocation scheme must not put excessively high overheads on any of the entities involved, such as the CA or the revocation repository, which will create undesirable service bottlenecks.

1.2.6 Concise yet Practical Formal Reasoning on PKI-based Protocols

Nowadays, it is common for many programs running on a host to interact with external entities using PKI-based operations. PKI-based protocols, including those used in secure program distribution and certificate revocation management, must be shown to be secure. Designing a correct protocol specification is however well recognized as a difficult task. Hence, a formal protocols analysis is necessary to establish the security of the protocols. Despite the availability of numerous formal techniques, one challenge is to devise techniques which can be handily utilized to verify real-world protocols by many protocol designers. One approach is by leveraging the popularity of widely-known techniques such as BAN Logic [25, 26]. BAN Logic however does not properly deal with the detailed aspects of PKI-based authentication, such as certificate processing, as commonly practiced nowadays. Given the ubiquity of PKI-based protocols and their importance to host security, there is a need to update such logic to be more concise yet remain practical to use.

1.3 Contributions

Based on the PPLC model and various security issues faced in securing it, we have proposed a number of schemes to help ensure more secure program environments. Most of the results here have been reported in the publications [146, 59, 173, 148, 147]. Appendix F gives a list of the author's published as well as submitted works throughout his Ph.D. candidature. The contributions of this thesis can be summarized as follows.

- Using our notion of *pseudo subtrace* construction, we demonstrate the security weakness of the Self-based IDS (Stide) [60, 140, 141], which compares the system call trace of a process with a set of k-grams (see Chapter 3). We then outline an efficient algorithm that transforms a previously detectable attack into the one which will pass the IDS detection. Using a number of real-world vulnerable programs, we show that mimicry attacks can be easily constructed with execution times of at most a few seconds, even when a relatively large window size is used. A variant model using a more precise *graph* profile representation can also be easily attacked. Based on this automated attack generation approach, we then propose a general framework which shows how the attack construction can be applied to various IDS models, including a Finite State Automaton (FSA) based IDS model [136]. This allows us to find out how computationally expensive it is to perform a search to craft attacks on the IDS. This result shows the feasibility of obtaining quantitative measurement on IDS resistance against targeted attacks.

- We also propose an extension to the Self-based IDS to make it more resistant against mimicry attacks by making use of the system call arguments and process credentials (see Chapter 4). The proposed data-flow technique can be easily combined with other system-call based IDS techniques. To avoid increasing the false positives, a user-supplied specification is used to abstract the arguments and credentials. This specification takes into account security semantics of the operations and their potential effects to system security, and highlights the occurrence of "dangerous" operations. With this simple extension, we show that the robustness of the IDS is increased. We demonstrate that the enhanced IDS provides resistance to previously successful attacks on the Self-based IDS and variety of common attack strategies. We also have some evidences that the increase in detection accuracy does not lead to more false positives. Unlike other proposed data-flow IDS techniques [83, 112, 155, 22], our IDS enhancement using system call arguments highlights the virtue of incorporating the host's *security model* into the IDS model to result in a more concise and robust IDS.

- In the area of software integrity protection, we demonstrate the practicality of a mandatory, in-kernel, pre-execution executable (binary) authentication mechanism

in Windows (see Chapter 5). We first analyze design factors for a binary integrity system and compare a variety of existing systems. We then present a prototype, BinAuth, which exemplifies a lightweight binary integrity scheme and its integration within an OS. The main contribution of this work is that we believe that it provides the first comprehensive infrastructure for trusted binaries in Windows. It provides mandatory authentication for all binaries in Windows, and goes beyond authenticated code in XP and Vista. This is significant given that much of the problems of security on Windows stems from a host's inability to distinguish between trusted and untrusted software. Our benchmarking shows that the overhead of the scheme can be quite low, around 2%, with a caching strategy. In addition, we also leverage the authentication with a Software_ID scheme which is useful for easier software management and automated vulnerability assessment.

- We propose a framework for machine-oriented vulnerability database geared towards automated vulnerability checking on a host (see Chapter 6). We developed a prototype sample scanner for Unix/Linux systems tailored for Red Hat Linux and FreeBSD, although the basic principles are applicable to other OSes. Our work was developed when vulnerability representation and automated processing were still not widely addressed by the security community. Our results are in line with present on-going standardizations efforts on vulnerability processing by the U.S. Government-sponsored organizations, such as CVE [105], OVAL [106], CPE [104] and SCAP [115]. Given these latest standardization efforts, we also address how these standardization results still need to be extended in order to realize the fully automated vulnerability processing as envisioned in our proposed framework.

- We also propose a family of lightweight, practical and inherently-distributed certificate revocation schemes called CREV, which are based on the recently available Extended-Validation (EV) certificate infrastructure (see Chapter 7). Our CREV schemes enhance CRS/NOVOMODO [100, 101] and OCSP [113] to support efficient revocation with near real-time timeliness guarantees (1–10 minutes). Using a realistic model derived from empirical data, we demonstrate that it is feasible to keep the overheads manageable on all the involved entities, while maintaining lightweight requirements on the verifier even under a timeliness guarantee of one minute. The scheme has an additional advantage in that status queries are protected from external parties including the CA, which may be important for privacy reasons. Another contribution is our development of a realistic model and a cost analysis framework for evaluating revocation schemes. We follow the approach in [91, 65] which makes use of empirical revocation data. Our framework incorporates a number of novel aspects, including: more accurate calculation of CRL data size over [91, 65] (which also works for fine-grained certificate generation and CRL is-

suance), a more realistic query model on revoked certificates, and the derivation of query probability on revoked and valid certificates.

- Finally, we propose various public-key related enhancements to BAN Logic [25, 26] to allow for more concise reasoning on PKI-based protocols (see Chapter 8). Our extension is along the lines of the work by Gaarder and Snekkenes [48] but better captures the current aspects of PKI. In particular, our extension redresses the reasoning on the goodness of private keys, and considers certificate revocation. We also address common pitfalls in public-key based protocol design due to insufficient attention placed on the "intended recipient" and "stated sender" of a message. Our extension makes the recipient and sender explicit, and these requirements are incorporated into the logic. As a result, our logic reduces the likelihood of allowing such flaws in the protocol and the corresponding formalism. In addition, our logic also deals with cases of (public-key) signed encrypted message and encrypted signed message. Examples on the usage of the logic are given. We also apply our logic to analyze a session establishment protocol proposed in CREV scheme.

1.4 Organization of the Thesis

The rest of the thesis is organized as follows. Chapter 2 gives background information on related problems addressed in this thesis. Chapter 3 presents our automated attack algorithm on the Self-based IDS and the framework for analyzing IDS resistance level. Chapter 4 proposes the IDS enhancement using system call argument and privilege abstraction. Chapter 5 outlines our lightweight, in-kernel software integrity protection scheme. In Chapter 6, we expound the automated vulnerability alert processing scheme. We present our lightweight and practical certificate revocation schemes in Chapter 7. Chapter 8 explains our extension to BAN Logic for reasoning with PKI-based protocols. Finally, Chapter 9 gives a summary on what have been achieved in this thesis and points out possible future work.

Chapter 2

Background

In this chapter, we give some background on the related aspects of system and network-based infrastructure that support secure program environments. This chapter can be skipped given the familiarity of the subjects, or its parts can just be referred to as necessary.

2.1 Intrusion Detection Systems

2.1.1 Overview and Motivation

Intrusion detection is defined as the process of monitoring events occurring in a computer system or network and analyzing them for signs of possible *incidents*, namely violations or imminent threats of violation of computer security policies, acceptable use policies, or standard security practices [130]. An Intrusion Detection System (IDS) is a device or software that automates the intrusion detection process. Intrusion Prevention System (IPS) has all the capabilities of an IDS and can also attempt to stop possible incidents. Prevention features in IPSs can usually be disabled by system administrators, causing them to function as IDSs. Since this thesis focuses on intrusion detection mechanisms, for ease of discussion, the term IDS is used throughout this thesis to refer to both IDS and IPS.

IDSs represent a system's second line of defense, which are aimed to provide additional layer of security protection. The adoption of IDSs is motivated by the following observations and trends [81], which have led many experts to belief that a computer system even with all its attack prevention measures will never be absolutely secure [19]:

1. Statistics show that the number of reported vulnerabilities has increased greatly in the past decade [29].
2. Meanwhile, the number of produced malware is continuously increasing [42].
3. Administrators are generally very slow in applying fixes to vulnerable systems [127].

4. There is an increasing concern on zero-day attacks to computer systems. Unreleased zero-day exploits even now have formed a "zero-day market" [103].

For an introduction to IDS, we recommend [72, 38].

2.1.2 IDS Classification

In this section, we give a very short overview of IDS with respect to the classification of IDSs. Comprehensive survey and taxonomy of IDSs can be found in [13, 138, 90].

IDSs are commonly classified by their *audit (event information) source location* and *detection method*. Based on the audit source location, the IDSs are classified into:

- **Host-based IDSs:** which deal with audit data generated on a single host.
- **Application-based IDSs:** which work on audit records produced by a particular application.
- **Network-based IDSs:** which monitor network traffic.

Based on the detection method, they can be broadly classified into:

- **Misuse detection IDSs:** which model known attacks using attack patterns (also called *signatures*), and detect them by means of pattern matching. The benefit of these IDSs is a high degree of detection accuracy, while their main drawback is the inability to identify new attacks.
- **Anomaly detection IDSs:** which define what is the "normal behavior" on the system (i.e. *normal profile*), and flags any deviations from normal as potential attacks. They thus operate based on the assumption that all anomalous activities are malicious [81]. Anomaly detection IDSs, also sometimes called *anomaly detectors*, are capable of detecting novel attacks. However, there are some challenges such as defining accurate "normal" behaviors and keeping low false positive rate. In fact, these challenges are not unique only to the anomaly detection IDSs, but represent known issues in the *anomaly detection problem* (see [31] for a comprehensive survey of the problem). Patcha and Park [121] gave a recent survey of anomaly detection techniques used specifically for intrusion detection purpose.

2.1.3 IDS Effectiveness Metrics

The effectiveness of an IDS is usually measured in terms of detection (true positive) and false positive rate. Functioning as a classifier on binary decision problem, an IDS determines examples as either positive (intrusive) or negative (non-intrusive). The decision can be represented in a structure known as a confusion matrix shown in Table 2.1.

The confusion matrix has the following four cases, and makes the four corresponding sets:

		Actual Condition	
		Intrusive Behavior	**Non-Intrusive Behavior**
Detection	**Intrusion Alarm**	*True Positive (TP)*	*False Positive (FP)*
	No Alarm	*False Negative (FN)*	*True Negative (TN)*

Table 2.1: Confusion matrix comparing actual intrusive condition and detection result.

- *True Positives (TP)*: refer to intrusive behaviors correctly determined as intrusion.
- *False Positives (FP)*: are non-intrusive behaviors incorrectly labeled as intrusion.
- *False Negatives (FN)*: represent a failure of an IDS to detect actual intrusions.
- *True Negatives (TN)*: when no intrusive behaviors have taken place and no alarm is raised by the IDS.

In a benchmark measurement scenario, based on the sizes of four defined sets, we can determine the following metrics [12] (let I and $\neg I$ denote intrusive and non-intrusive behaviors respectively, and let A and $\neg A$ denote the presence or absence of an alarm respectively):

- *True positive (detection) rate*: is defined by the conditional probability $P(A|I) = \frac{|TP|}{|TP| + |FN|}$, where $|\ |$ denotes the size of a set.
- *False positive rate*: is defined by the probability $P(A|\neg I) = \frac{|FP|}{|FP| + |TN|}$.
- *False negative rate*: is $P(\neg A|I) = 1 - P(A|I) = \frac{|FN|}{|TP| + |FN|}$.
- *True negative rate*: is $P(\neg A|\neg I) = 1 - P(A|\neg I) = \frac{|TN|}{|FP| + |TN|}$.

2.1.4 IDS and Alert Correlation

Alert correlation has recently attracted the attention of the IDS community as an extension to intrusion detection. It is defined as a process that takes as input the alerts produced by one or more IDS sensors, and provides as output a more succinct and high-level view of occurring or attempted intrusions. See [81] for a survey on alert correlation.

2.2 System-Call Monitoring IDSs: Self-based IDS, Attacks, and Related IDS Models

In this part, we provide some background on system-call monitoring IDSs, particularly the Self-based IDS which is the focus of Chapters 3 and 4. A recent paper by Forrest et al. [47] gives a comprehensive survey on the development of system-call monitoring IDSs as inspired by the Self-based IDS.

2.2.1 Self-based IDS

Inspired by how the natural immune system distinguishes "self" from "other", a seminal work [60] proposed a host-based anomaly detection IDS based on an observation that *short sequences* generated from a program's system call traces can serve as a stable indicator of the program's behavior. Monitoring system calls has been recognized as an effective mechanism to observe what a program really does. System calls are the gateway to privileged kernel operations. Thus, by monitoring and restricting them, a running program can be observed and (if necessary) be prevented from violating the host's security policy.[1] The works [140, 141] expanded the IDS with an automated response to a detected intrusion. For ease of reference, in this thesis we will call this type of IDS a *Self-based IDS*.

The Self-based IDS takes an unparameterized system call trace of a program as its input. A system call trace can be viewed simply as a string where the alphabet consists of all the possible system calls in the underlying OS (∼200 for Unix/Linux systems). The IDS defines normal behavior of a program in terms of short k-grams of system calls. A k-gram is a sequence of system calls with length k. Conceptually, we define a small fixed size (k) window and then "slide" it over each normal trace. The parameter k is thus usually known as the *(sliding) window size*. Every unique system calls within the sliding window, i.e. k-gram, is recorded into the normal profile database. Sliding-window sizes of 6–8 are commonly used in practice [60, 140, 169].

Previous works on Self-based IDS make use of a few different techniques for representing a normal database entry, namely *full sequence*, *(forward) lookahead pair* and *backward lookahead pair* technique. The Self-based IDS using full sequence representation is also known as *sequence time-delay embedding* (Stide). It was first proposed in [60], and is later analyzed in a widely cited performance comparison work [169]. Stide model has been widely used for performance analysis mainly due to its better discrimination power than the lookahead pair methods. However, a recent work comparing Stide and the backward lookahead pair method [67] shows that the latter exhibits lower false positives. Notwithstanding this, in our attack construction in this thesis, we mainly consider Stide since it is theoretically a more resistant model against mimicry attacks [67, 169].

Experiments were conducted in [60] to evaluate the effectiveness of the Self-based IDS. In [169], Warrender et al. reported the comparison results of the Self-based IDS and several system-call analyzing IDSs including an IDS based on data mining and Hidden Markov Model. Although no definite conclusion can be made, the results show that the Self-based IDS performs reasonable well despite its relative simplicity.

[1] This system-call based monitoring approach cannot detect intrusions which *do not* invoke system calls. Hence, this approach performs simpler measures in ensuring the intrusion-free (unaltered execution) property (see Section 1.1) of a running program. In practice, however, security-relevant interactions typically take the form of system calls [167].

2.2.2 Mimicry Attacks on Self-based IDS

Despite the good experimental results in [60, 169], the security of the Self-based IDS did not receive much attention until later when Wagner and Soto [167] and Tan et al. [152] independently published their security analysis results on the IDS.

Wagner and Soto [167] used Finite State Automaton (FSA) as a framework for studying and evaluating mimicry attacks. They showed that a mimicry attack is possible because additional system calls which behave like no-ops can be inserted into the original attack trace so that the resulting trace is accepted by the automaton of the IDS model. Independently, Tan et al. [152] showed attack construction as a process of moving an attack sequence into the IDS detection's *"blind region"* through successive attack modifications. The focus in these two works is to demonstrate the feasibility of mimicry attacks. However, they do not take a detailed look at designing and constructing automated attacks.

In a later work, Gao et al. [50] performed a study of the "black-box" Self-based IDS and also several "gray-box" IDSs such as [136, 45].[2] They investigated mimicry attacks with window sizes of up to 6 and showed the existence of mimicry attacks across the methods and the window sizes studied. However, they did not go into the details of attack generation. Later works, including the more recent mimicry attack construction techniques [80, 53], are surveyed in Section 3.1 when we discuss our own automated mimicry attack construction.

2.2.3 Improved System-Call based IDSs

Due to possible attacks on the basic Self-based IDS, a variety of improved IDS models have been proposed. Within the context of black-box and gray-box anomaly detectors, there are two main approaches, namely:

- **Incorporation of additional execution context**. Two main examples of this approach are [136, 45]. In [136], Sekar et al. proposed the use of FSA-based IDS which is generated by observing both system calls and program points during the normal program executions. A *program point* is the Program Counter (PC) at the point from where a system call is made. A stack traversal mechanism is used to recover the PC within the program segment where a system call is actually invoked. Each distinct PC is made as a state, whereas system calls are used as the labels

[2] We follow the terminology in [50] which classifies various system-call monitoring anomaly detectors into "black-box", "gray-box" and "white-box" detectors. Black-box detectors (e.g. [60, 140]) do not acquire any additional information other than the system call number and arguments. In contrast, white-box detectors (e.g. [166, 44]) examine all available pieces of information including those acquired by statically analyzing (and potentially modifying) the source or the binary codes. Gray-box approaches (e.g. [136, 45]) lie in between: the anomaly detector does not utilize static analysis of the program, but does extract additional runtime information.

for transitions. The FSA can then be used to monitor the program execution in an online fashion. If an error occurs while performing the stack traversal mechanism, or if a state or transition does not exist, there may be an anomaly. Since the FSA model is deterministic, the efficiency is high. It, however, allows the possibility of false positives as some legal transitions or states may never occur during training. In addition, it also suffers from the impossible path problem [45]. Feng et al. [45] then proposed the use of call stack information obtained through runtime monitoring in addition to the program's system call trace. Later works [50, 80], however, showed that attacks which will escape the detection of this IDS model can still be constructed.

- **Data-flow approach**. Examples of IDS models incorporating data-flow analysis which can be used to improve the Self-based IDS against mimicry attacks are [83, 112, 155, 22]. We discuss more about these models when we present our work on improving the Self-based IDS in Chapter 4.

A survey paper [47] also mentioned other works making use of static analysis and other observables, which are beyond the scope of this thesis.

2.3 Software Integrity Protection

Chapter 5 addresses the problem of establishing the integrity of a piece of software (i.e. executable codes). For ease of reference in this thesis, we refer to an executable code stored in the file system as a *binary*. The goal of a software integrity protection system is to ensure that an executed binary only comes from trusted software developers, and that it is executed in the correct execution context. Below, we define the authentication goals, and survey the existing systems aimed to provide these goals.

2.3.1 The Executable Integrity Problem

Many of the system security attacks stem from the fact that distrusted code is executed on the system. A modern OS has numerous built-in security measures designed to prevent illegal operations on a system, including illegitimate addition or modification of executable code on the file system. Due to various vulnerabilities, an OS component or an application program however can be attacked. Once an attacker succeeds in compromising a system, the next step is usually to install/modify executable code(s) on the victim's file system. This could be done for several reasons: to install a backdoor, to plant a spyware, or to use as a step for subsequent privilege-escalation attacks. A mechanism to deal with illegal addition or modification of executables on a system is thus required.

This would also help prevent social engineering attacks which attempt to trick a user to install a software package that illegitimately replaces important system libraries.

Besides protecting application programs, binary authentication is also beneficial to protect the OS against illegal loading of malicious drivers into its kernel.

2.3.2 Overview of Existing Authentication Systems

Although the concept of binary integrity authentication is not new, it is not commonly incorporated as a standard in-kernel mechanism in popular commercial off-the-shelf OSes, such as Windows and Solaris. Below, we will briefly mention some existing systems as well as other security approaches related to establishing software integrity protection.

Tripwire [74] is one of the first to deal with file integrity protection. However, Tripwire is a user-mode application program. It checks file integrity in an off-line manner, and does not provide any mandatory form of integrity checking.

There exist a number of kernel-level binary authentication implementations. These are mainly for Unix, such as DigSig [8], Trojanproof [170] and SignedExec [162]. They modify the Unix kernel to verify a binary's digital signature before executing the binary. For efficiency, DigSig employs a caching mechanism to avoid checking binaries which have already been verified.

Partly motivated by the emergence of mobile code (e.g. Java applets), some OSes extend the concept of signed code to standard binary types. They check the validity of a binary's signature particularly if it is downloaded from an untrusted zone like the Internet. Windows has a technology called Authenticode [55]. In Windows XP and earlier versions, Authenticode alerts the users of the results of signature verification under a few situations. However, it is not mandatory, and can be bypassed. Windows Vista User Account Control (UAC) adds binary signature checks in some special cases, but it only deals with EXE binaries. It is also limited to privilege escalation situations. One common drawback of existing Windows mechanisms is that they do not authenticate the binary's pathname. Moreover, they always perform digital signature verification when verifying a protected binary. Besides the incurred performance overhead, a certificate revocation service is thus always required. Hence, they may be too heavy for an online (pre-execution) program integrity checking mechanism, particularly on lightweight computers.

There is also the Trusted Computing initiative by the Trusted Computing Group (TCG) [160]. The Trusted Platform Module (TPM), promoted by the TCG, is a certified hardware in which all the basic trusted operations and the cryptographic functions are securely handled. It thus constitutes the "root of trust" aimed to provide an uninterrupted "chain of trust" which starts at the system's startup up to the application's executions. TPM can be employed, among others, to facilitate whole-disk encryption (e.g. BitLocker) and secure startup. We observe that TPM indeed provides a higher

assurance level of system security. An authentication system like our proposed BinAuth scheme can take advantage of TPM by having the latter securely store keys and authentication data. In addition, TPM can facilitate a secure startup service until the OS kernel (with its program integrity protection module) takes control to protect subsequent operation of the application programs.

The concept of layered integrity protection, which aims to ensure the integrity of the overall system starting from the lowest level (hardware) to the highest level (application), was analyzed among others by Arbaugh [9]. In the layered model described in [9, page 54], an in-kernel authentication scheme deals with integrity assurance at level 5 (user applications) and part of level 4 (some components of the OS, e.g. kernel drivers).

2.3.3 Authentication Issues in Microsoft Windows

Most works on binary integrity authentication are on Unix/Linux [8, 170, 162]. The problem of malware is however more acute in Microsoft Windows. Windows is perhaps the most commonly exploited commercial OS. This is partly due to its sheer complexity, which leads to a high number of security flaws. Below, we briefly mention the complexities and special problems of Windows which make it more difficult to implement binary authentication than in other OSes such as Unix. This is to highlight issues which need to be taken into consideration when establishing a binary authentication system on Windows.

Windows NT (including later Windows versions based on it such as Server 2000, XP, Server 2003, and Vista) is a microkernel-like OS. Programs are usually written for the Win32 API, but these are decomposed into microkernel operations. Windows is closed source — only the Win32 API is documented but not the microkernel API. As a result, it is not possible for an enhancement work related to Windows kernel infrastructure to make any guarantees on the completeness of the security mechanisms.

Some specific issues in Windows related to protecting the binary integrity are as follows [59, 173].

- **Proliferation of binary types:** It is not sufficient to ensure the integrity of `EXE` files alone. In Windows, binaries can have any file name extension, or even no extension. Some of the most common extensions include `EXE` (regular executables), `DLL` (dynamic linked libraries), `OCX` (ActiveX controls), `SYS` (drivers), `DRV` (drivers) and `CPL` (control panel applets). Unlike in Unix, binaries cannot be distinguished by an execution flag.
- **Complex process execution and DLL loading:** A process is created using `CreateProcess()` which is a Win32 library function. To load a DLL, a process uses `LoadLibrary()`. However, these two operations are broken up into several operations at the native API level. Hence, it is more complex to incorporate a

mandatory authentication in Windows. We need to properly intercept the right API(s) with correctly intended operation semantics with respect to Windows' behavior.

Compared to other open platforms, Windows potentially also makes the issue of locating vulnerable software components more complicated. A great deal of binaries created by Microsoft contain an internal file version, which is stored as the file's meta data. The Windows update process, however, does not inform the user which files have been modified. Furthermore, the meta data of the modified file might remain unchanged. Thus, merely by inspecting the meta data of a binary, one cannot ensure whether the binary is still affected by a particular vulnerability. Nonetheless, a software naming scheme, which associates binaries with their respective unique names, can simplify the software vulnerability management. This issue is addressed later in Section 5.6.

2.4 Managing Host Vulnerabilities

2.4.1 The Problem of Host Vulnerabilities

A *vulnerability* is defined as a weakness in a system that allows the use of a product beyond its design intent with an adverse effect on the software, system, or data [46]. Security loopholes found in programs and OS thus lead to vulnerabilities in the system.

In recent years, the number of security vulnerabilities discovered in computer systems has increased explosively [29]. To keep track of known vulnerabilities, system administrators usually rely on vulnerability alert repositories/databases[3], such as CERT Coordination Centre Vulnerability Notes Database (now hosted by the US-CERT [161]), Securityfocus BugTraq [135] and Open Source Vulnerability Database [158]. Unfortunately with the observed rate of alerts (e.g. 5,500 vulnerabilities reported in 2002), it has been estimated that it would take more than 200 days for an administrator to digest all the alerts published in that year alone [30, page 41]. Hence, a mechanism which requires human intervention to understand and deal with such alerts is too time consuming, and more importantly is too slow to respond to vulnerabilities in this day and age.

This situation, we argue, has contributed quite significantly to the noticed trend in the *Vulnerability Exploit Cycle* [86] shown in Figure 2.1. The Y-axis in the graph represents the number of incidents for a given vulnerability. It quite clearly shows the existence of *significant time lag* between the release of a vulnerability report and the decrease of incidents following corrective measures by users/Administrators.

[3]Following the common practice, we refer to a vulnerability alert repository as vulnerability database although the alert entries are narrative and not structured using a well-defined database schema, e.g. relational model.

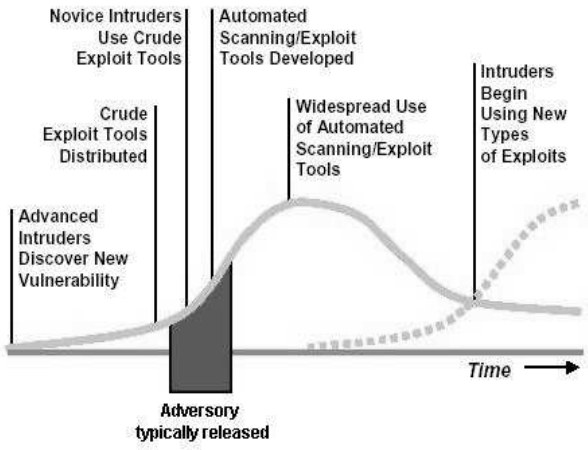


Figure 2.1: Vulnerability Exploit Cycle (from CERT Coordination Center [86]).

2.4.2 Vulnerability Assessment and Self-based IDS

A *vulnerability assessment tool* (also known as *vulnerability scanner*) tests to determine whether a network or host, which is generally called target, is vulnerable to a set of known attacks [46]. For each vulnerability, it performs a *vulnerability check*, which is a particular set of items to check on a target that may reveal the presence of the vulnerability [46]. As acknowledged by many including [15], vulnerability analysis is a very powerful security management technique. However, it also has its own limitations, and is suitable only as a *complement* to using an IDS, *not as* a replacement. It thus works together with an IDS in protecting the PPLC as depicted in Section 1.1. In [15], Bace views that a vulnerability assessment can be treated as a special case of intrusion detection process. To better highlight the differences between a vulnerability assessment tool and the Self-based IDS (as an anomaly detector), we compile the comparison of the two systems in Table 2.2.

Vulnerability scanners can be broadly classified into two categories:

1. Network-based scanners: which probe a target machine remotely to find vulnerabilities. Widely known examples are SATAN [134] and Nessus [117][4]. These scanners can help administrators automate the process of testing target systems for known vulnerabilities that can be exploited via the network. In its typical operation, Nessus begins by performing a port scan to determine which ports are open on the target, and then tries various exploits on the open ports.
2. Host-based scanners: which are installed and executed on the target host itself, such as COPS [43] or Ferret [137]. The host-based scanners have an advantage over network-based scanners as they can directly access the host's configuration

[4]Nessus is traditionally and still mainly a network-based scanner. Recent development however also equips it with an option to perform host-based assessments. Nessus makes use of *network-based access* using system credentials rather than a software agent approach [39]. It logs on to Unix systems via the SSH protocol and to Windows systems via an NTLM network API.

Aspect	Vulnerability Assessment Tool	Self-based IDS
Goal	To detect vulnerable programs in a system	To detect whether an intrusion is taking place on a program
Scope of detection	Whole system (OS and application programs)	A program of interest
Information examined	System state, i.e. installed programs and environmental settings	System call trace invoked by the observed program
Reference data	Vulnerability alert entries from vulnerability database center(s)	Normal program profile (database of k-grams)
Type of analysis	Misuse detection	Anomaly detection
Timing	Interval-based (each execution makes a snapshot of the system's state)	Whole program execution
Output	To report whether there exist vulnerable programs in the system	To raise an alarm when anomalous behavior is observed
Monitoring agent	System information collector as part of the vulnerability scanner	System call interposition (reference monitor)
Main strength	Exhaustively test a system for large numbers of known vulnerabilities	Able to detect novel attacks or anomalous behaviors
Main drawback	Offer no protection for attacks in between scanning times and no ability to detect novel attacks	Must continuously observe the program and inspect its activity

Table 2.2: The comparison between vulnerability assessment tool and the Self-based IDS.

information and various running services. They can also work together with a "scanning manager", which is deployed within the same administrative network domain. In this way, the scanners can be made as lightweight as possible. As such, we can mitigate a main drawback of host-based scanners which require installations or regular updates of the scanners on every target machine.

Our work reported later in Chapter 6 focuses on developing a host-based scanner to automatically process host vulnerability alerts.

2.5 PKI and Certificate Revocation

Public Key Infrastructure (PKI) is defined in [68] as "the infrastructure able to support the management of public keys able to support authentication, encryption, integrity or non-repudiation services". The dominant PKI standards used are the ITU-T X.500 standards, in particular X.509 [68]. X.509 has also been adopted by the IETF (Internet Engineering Task Force) for use in the Internet [36]. PKIs have been developed and deployed to facilitate secure communication and transactions over insecure public networks such as the Internet. Electronic commerce and secure server-access applications that use Transport Layer Security (TLS) or Secure Sockets Layer (SSL) [40] do rely on PKI. Verification of network-based (signed) software distribution and mobile code execution also critically depend on a timely and efficient PKI service.

In order for a principal to trust a public key belonging to another principal, the public

key must be obtained from a trusted source commonly known as CA. A CA certifies a public key by issuing a digitally sign certificate, which binds the public key to the entity holding the corresponding private key. An issued certificate is expected to be in use for its entire validity period. However, it can become invalid prior to its expiration due to several reasons, such as change of name, change of association between the subject and the CA (e.g. the subject is no longer with an organization), and compromise or suspected compromise of the corresponding private key. Under such circumstances, the CA needs to revoke the certificate.

From technical viewpoint, certificate revocation is probably the most challenging task of certificate management [88]. It thus deserves a special attention, and needs to be satisfactorily addressed so that the PKI service can be reliably available to host systems.

2.5.1 Issues in Certificate Revocation

In X.509-based PKI, certificate revocation is mainly supported by a mechanism called Certificate Revocation List (CRL) [36]. However, CRL is widely perceived to be costly and is attributed as one of the main impediments to the successful global PKI deployment [88, 58]. Among others, CRL suffers from a high bandwidth requirement and (generally) a low timeliness guarantee on the part of the verifier.

Several alternatives, such as Online Certificate Status Protocol (OCSP) [113], have been proposed. OCSP offers a potentially real-time recency, but the CA needs to respond to any incoming query in real time and reply with a signed message. As such, it imposes significantly higher computation and network requirements on the CA. CA/Browser Forum, for instance, mandates OCSP support for its premium certificates only starting from 2012 [27].

To secure software distribution and programs' interaction with external hosts (e.g. using TLS/SSL), certificate revocation service must provide a *near real-time* timeliness guarantee and enable a fast computation for the verifier. This is particularly important given the proliferation of lightweight computers and mobile devices. In addition, the revocation scheme must not put excessively high overheads on any of the entities involved. Otherwise, undesirable service bottlenecks will be formed. Chapter 7 proposes a family of lightweight and practical certificate revocation schemes which provide a near real-time timeliness guarantee.

2.5.2 Survey of Existing Certificate Revocation Systems

Here, we look at various existing revocation schemes, and examine the potential issues with them. Throughout this thesis, we will refer to the subject of a certificate as a *principal*, and the party accepting/verifying certificates as a *verifier*. We also adopt the term *Certificate Status Information (CSI)* [66] to denote information pertinent to

the validity of a certificate. Hence, CSI encompasses CRL data, OCSP messages and hash tokens in CRS/NOVOMODO. We refer to *Certificate Management Ancillary Entity (CMAE)* as a generic term for a *repository* or *directory*, from which a verifier may obtain the CSI.[5] Also, we follow the definition of *revocation timeliness guarantee* [3] as the time interval between when a CA makes a revocation record and when it makes that information available to the verifiers.

Standardized in X.509 [68] and also profiled in the IETF [36], CRL is the most widely supported revocation scheme. Rivest [128] criticizes CRL by elaborating its several important drawbacks. McDaniel and Rubin [97] however responded to the criticism by pointing out that CRL still can provide a useful and efficient service in some environments. For many typical online transaction environments, however, CRL is widely known to have serious shortcomings. Firstly, the CRL may eventually grow to a cumbersome size in very large PKIs, e.g. ∼750KB in Verisign's [131] and ∼40MB in the U.S. DoD's [118]. Secondly, the downloaded CRLs may be mostly redundant given that more than 90% of the information is irrelevant to the verifiers [131]. Lastly, CRL does not offer adequate timely revocation guarantees. It is common for a CA to update its CRL only *daily*, as suggested in [164], although the CA may have a service for a shorter window period presumably with a higher charge.

There are several methods for improving the basic CRL mechanism, such as CRL Distribution Points, Delta CRLs, and Indirect CRLs (see [3] for a survey). However, all these improved schemes still put the same requirement on the verifier to obtain a complete revocation list, which includes the unrelated entries.

OCSP [113] was proposed to provide a more timely certificate status checking. An OCSP Responder is required to return the status information about a specific certificate in a digitally signed response. Since OCSP is an online service, it necessitate a prompt and reliable OCSP Responder system with a high level of security. It basically shifts the demands from network bandwidth in CRL to the Responder's processing power [131]. The high requirement put on the Responder to promptly generate signed replies may however be an issue [88]. A report on day-to-day operation of the EuroPKI infrastructure [85] highlighted the signature operations as the Responder's main bottleneck. Furthermore, the Responder's throughput is always bounded by the maximum number of signatures per second that it is able to perform.

Several modifications have been proposed on OCSP. H-OCSP [111] is as an improvement over OCSP, which allows a verifier to verify the validity of a pre-produced response beyond its documented life time (as indicated by `thisUpdate` and `nextUpdate` fields in OCSP Response). The scheme uses a hash-chaining technique so that a H-OCSP Response can later be updated with reduced information and processing needs. However,

[5] A CMAE may be trusted or non-trusted depending on the revocation scheme.

schemes like H-OCSP requires the CA to cater for a potentially large number of verifiers. Since the CA needs to maintain a hash-chain for each verifier, its bandwidth and storage requirements remain high.

Merkle Based Server (MBS)-OCSP [20] is a further improvement over H-OCSP, which employs Merkle Hash Tree (MHT) for faster verification by the verifier. In MBS-OCSP, a Responder associates a MHT with an OCSP Response, and stores in its local database the signed OCSP Response together with the associated MHT. We note that the use of MHT in MBS-OCSP is essentially the same as that suggested by Naor and Nissim [114]. The enhancement, however, comes with at the expense of larger messages to be downloaded by the verifier and extra storage on the Responder for all the MHTs.

Another approach, sometimes called "trusted directories", includes Certificate Revocation Scheme (CRS) [100] and NOVOMODO [101]. CRS/NOVOMODO, which we improve upon in our schemes in Chapter 7, are reviewed in Section 7.2.2.

Aiello et al. [5] proposed an improvement to CRS called "Hierarchical Scheme". It is aimed at reducing the CA-to-CMAE communication while still maintaining the low query communication. The improvement however comes at the price of a significant increase in the certificate transmission costs due to the increase in certificate size.

Kocher [75] proposed Certificate Revocation Tree (CRT) which employs MHT. In CRT, the CA breaks its revoked certificate list into a series of ranges that describes the certificate statuses. The lower end of each range contains the unique serial number of a revoked certificate. The ranges are then packaged as leaf nodes. A tree is built on these leaves by hashing two sibling values to obtain their parent node. The scheme however suffers from a high computational cost needed to update the tree.

Naor and Nissim [114] extended CRT by using a more suitable data structure, 2-3 tree rather than a binary hash tree. In 2-3 tree, insertion and deletion on an element only affects a few interior points. Thus, there is no need to recalculate the entire tree.

2.5.3 Extended-Validation (EV) Certificates

In Chapter 7, our proposed scheme makes use of the Extended-Validation Certificate infrastructure as the basis for a lightweight and practical revocation scheme. We give some background on EV Certificate (EVC) here.

Recent attacks, such as certification-chaining attack [94], homograph phishing [49, 62, 94], and man-in-the-middle based SSLstrip attack [94], have shown weaknesses in the usage of the standard X.509 certificate. As a response, the EVC initiative was launched by the CA/Browser Forum [27] to provide stronger assurance on certificates. This is achieved through a stricter issuance process based on the EV Guidelines [27]. The Guidelines standardize the identification process for EVCs, so that those issued by different CAs are still of the same quality with respect to identification. In addition,

the Guidelines specify more rigorous checkings by a CA before issuing an EVC, which include:

- verifying the legal, physical and operational existence of the entity;
- verifying that the entity has an exclusive right over the specified domain name;
- verifying the identity and authority of the individuals acting for the website owner, and that documents pertaining to legal obligations are signed by an authorized officer.

There have been a number of studies conducted to ascertain the usability and effectiveness of EVC and its purported benefits, particularly in securing SSL transactions [70, 156, 139]. These studies however focused mostly on the user's perception of EVC based on the perceived browser's interface. Jackson and Barth [69] evaluated the security of EVC with respect to the browser's security policy, and reported a possible vulnerability in its usage by the browser. We however view the issue more as a loophole in the browser's SSL usage of EVC rather than the flaw of EVC infrastructure. It is worth nothing that despite the suggested conclusions, all of the studies acknowledge the value of EVC in providing stronger protections for secure transactions.

2.6 Formal Protocol Verification and BAN Logic

Designing a correct protocol specification which satisfies certain security properties is well recognized as a non-trivial task. Many logics and formal techniques have been proposed for verifying cryptographic protocols. Among various authentication logics, Burrows-Abadi-Needham (BAN) Logic [25, 26] is one of the best known and most widely used [98, 133, 99]. One reason for its popularity is that BAN Logic is comparatively easy to use. As pointed out by Meadows [98, 99], BAN Logic's intentional avoidance of many advanced features makes it a simple and straightforward logic that is easy to apply yet of substantial use for detecting flaws. This may well explain the constant appearance of publications applying BAN Logic even till now [4, 23, 143, 175, 34, 32], with application domains as diverse as wireless network [175], mobile communication [32], and voting system [143].

2.6.1 Overview of BAN Logic

BAN Logic [25, 26] is a modal-sorted logic constructed on several sorts of objects: principals, keys, messages and well-formed formulae. In BAN Logic notation, the symbols P, Q and R are usually used to denote principals; whereas X and Y are statements; and K is an encryption key. *Predicate constructs* are used to interpret organized objects into well-formed formulae. BAN Logic defines the following constructs:

$P \models X$: P **believes** X. P may act as though X is true.

$P \triangleleft X$: P **sees** X. Someone has sent a message containing X to principal P, who can read and repeat X (possibly after doing some decryption).

$P \mid\sim X$: P **once said** X. P at some time sent a message including X. It is not known whether the message was sent long ago or during the current run of the protocol, but it is known that P believed X when he sent the message.

$P \Rightarrow X$: P **has jurisdiction over** X. P is an authority on X and should be trusted on this matter.

$\sharp(X)$: X **is fresh**. X has not been sent in a message at any time before the current run of the protocol.

$\{X\}_{K_{PQ}}$: X **encrypted with a secret key** K_{PQ}.

$P \overset{K_{PQ}}{\longleftrightarrow} Q$: P and Q **may use a secret key** K_{PQ}. The key K_{PQ} is good, i.e. it will never be discovered by any principal except P or Q, or a principal trusted by either P or Q.

The only propositional connective used in BAN Logic is conjunction, which is denoted by a comma. A set of inference rules, also called logical postulates, are used to reason on well-formed formulae. Appendix C lists all the rules of BAN Logic which are relevant to our discussion in this thesis.

BAN Logic has been criticized by many largely for its idealization process. It lacks a systematic way of transforming a protocol into the form that will be used by the logic. Additional criticisms were also pointed out in [116, 54, 93]. Fortunately though, along with these criticisms, suggestions for improvement as well as precautions on the improper usage of BAN Logic were made [48, 163, 73]. [163] attributes inaccurate idealization rather than the logic itself to be the root of many problems associated with BAN Logic.

For more discussion on BAN Logic, see [56, 98]. Syverson and Cervesato [151] give a tutorial on BAN Logic, and put it within a broader context of authentication logics.

2.6.2 Issues on BAN Logic Application to PKI-based Protocols

Despite being useful and widely applied, BAN Logic gives a rather simplified treatment of public-key authentication processing. It does not deal with deeper aspects of public key authentication such as certificate processing. This is presumably because PKI was not well established when the logic was designed. The situation is now very different since PKI becomes common.

There exist some works such as [4, 48, 144] which attempted to extend BAN Logic to reason better with public key authentication. We find that the extension proposed in [48] does improve the expressiveness of BAN Logic while keeping the logic's secret-key aspects intact for easy application. However, it still falls short in capturing many important concepts and practices of the modern PKI usage.

Chapter 3

Self-Based IDS: A Security Analysis and Automated Attack Construction

Section 2.2.1 gives an overview of the Self-based IDS [60, 140, 141], which compares an unparameterized system call trace of a process against the normal profile using k-grams of system calls. While the IDS have been shown to be effective in detecting intrusions [60, 169], it can be susceptible to evasion or mimicry attacks [167, 152]. The attacks disguise an attack trace so that it appears "normal" to the IDS. Our aim in this chapter is to systematically investigate the susceptibility of the Self-based IDS to mimicry attacks. In particular, we focus on deriving *automated* and *practical* attack constructions, and whether changing parameter(s) and normal profile representation of the Self-based IDS can help prevent such attacks. Here we assume a Self-based IDS which makes use of the full k-grams, an IDS model sometimes referred to as Stide [169][1], instead of the lookahead pair representations.

In the discussion to follow, we present a branch-and-bound algorithm for automatically constructing the *shortest mimicry attack trace* on the Self-based IDS (Stide) and its variant where the normal profile is represented as a *graph* of k-grams. This variant (defined later in Section 3.2.3) is a more precise model since it records a temporal relationship between two consecutive k-grams generated from the normal trace. We evaluate these two models under two attack scenarios of *trojan attack* and *injection attack*. Our experimental results show that using sliding-window sizes larger than what is commonly used (6–8) and even using a more precise graph profile representation cannot prevent mimicry attacks on both attack scenarios. Furthermore, the execution times of the at-

[1]Stide (sequence time-delay embedding) makes use of a more precise representation compared to the lookahead pair counterparts. As a result, it gives better discrimination in detecting potential anomalous behaviors, and is theoretically more resistant against mimicry attacks [67, 169].

tack construction on several sample vulnerable programs are at most a few seconds, even for the relatively large sliding-window sizes (k=9–11). These running times thus show the practicality and efficiency of our attack construction algorithm.

In addition to producing the mimicry attack trace, our construction technique also provides us with useful information on how computationally expensive it is to perform a search to craft attacks on an IDS. Hence, it provides a systematic method to effectively measure the *resistance level* of the IDS against targeted attacks. Based on this observation, we generalize the attack construction method into a framework for measuring IDS resistance against attacks using a notion of *"attack space"*. We show the generality of the framework by applying it to other system-call based IDS models. Parts of the results on our mimicry attack construction on the Self-based IDS have been reported in [146].

The remainder of this chapter is organized as follows. Section 3.1 discusses related works on mimicry attacks. We give an algorithm for constructing mimicry attacks in Section 3.2. Section 3.3 presents the results of our experiments on automatically attacking the Self-based IDS and its variant. We give discussions on the results in Section 3.4. The generalization of an attack into a framework for measuring the robustness of IDS is explained in Section 3.5. Finally, Section 3.6 gives a chapter summary.

3.1 Motivation and Limitations of Existing Works

Wagner and Soto [167] showed how the Self-based IDS which takes a non-parameterized system call trace can be vulnerable against mimicry attacks. They achieved this by inserting system calls which semantically behave like "no-ops", i.e. dummy operations that do not change any system effects, into the attack trace. In their work, they model the detection mechanism of the Self-based IDS as a Finite State Automaton (FSA). The attack is also defined by a regular language. By adding no-ops, an attack is thus transformed into one which falls into the language modeled by the FSA of the IDS. Although the work contributed to the establishment of theoretical framework for the Self-based IDS from FSA viewpoint, the question of whether such attacks are practical is not fully answered. In our work, we address the practical construction of attacks and its efficiency issue.

Independently, Tan et al. [152] showed that it is possible to modify a foreign subtrace to be accepted by the Self-based IDS by taking advantage of the limited sliding-window size. They gave a hint on how this process could work involving an example with a sliding-window of length two. Their work viewed the construction problem as how to move an attack sequence into the IDS detection's *"blind region"* through *successive attack modifications*. However the work is primarily descriptive, and does not give a systematic procedure for attacking the IDS. A later paper by Gao et al. [50] showed some experi-

mental results of attack construction on the Self-based IDS (Stide), which is a black-box detector, and several gray-box detectors [136, 45]. They investigated mimicry attacks on the IDSs with window sizes of up to 6, and showed the existence of mimicry attacks across the methods and the window sizes studied. But there was also no mention in the work on how the attack construction procedure was actually performed, which is the main focus of this chapter. Furthermore, we also consider mimicry attack construction algorithm in both trojan and injection attack scenarios.

Based on these earlier works, we can conclude that although mimicry attack on the Self-based IDS is known, and examples of turning an attack trace into a stealthy one do exist, it is however not completely clear how to generate a *practical* attack, which is fully automated and efficient in terms of the construction time. Additionally, we are also interested to find out whether varying parameter(s) and the normal profile representation of the the Self-based IDS can help prevent mimicry attacks.

Other related works on mimicry attack construction are [80, 53]. Mimicry attack on gray-box detectors [136, 45] requires forging the call stack and also temporarily relinquishing control to the victim application before regaining it for the next system call execution. Gao et al. [50] suggested a technique to forge the call stack and transfer the execution control to the injected code. To regain control, they suggested the modification of any code pointer following the system call so that it points back to the attack code. The work [50] shows the feasibility of this technique on a small example program, but manual development of mimicry attacks for realistic programs based on the technique poses some challenges. Kruegel et al. [80] addressed this challenge with a novel technique based on symbolic code execution on a program's binary to automate the steps needed for regaining control. Using example programs and applications, they showed that about 90% of the times, control could be successfully returned to the attack code. This result highlights the susceptibility of gray-box detectors against mimicry attacks. The technique in theory reduces the difficulty level of constructing mimicry attacks on gray-box detectors into that on black-box ones (e.g. Self-based IDS). The problem now becomes that of creating a "stealthy" attack trace interspersed with no-ops, the process which our attack construction algorithm addresses.

Giffin et al. [53] generated mimicry attack sequences using techniques based on model checking. The input to the model checker includes a specification of the OS model, the program model in the form of Push Down Automaton (PDA), and a specification that characterizes "unsafe" OS states. By using the model checking, their technique automatically finds system call sequences, together with the necessary arguments, which are allowed as a valid execution by the PDA that however produce unsafe OS state(s). However, the work has its own limitations too. Firstly, although the system does not require an initial exploit (or *basic attack trace* in our terminology as discussed in Section 3.2.5),

it however requires a manually-specified model of the OS and its unsafe configurations. Secondly, the "proofs" of detection hold only *with respect to* the developed OS abstraction and may not hold in the actual OS implementation. Thirdly, depending on the model-checker used, a reported attack sequence may not be the shortest attack trace possible. Lastly, the sample generated attack sequences given in [53] do not attempt to reach the `exit()` system call in order to terminate the attack gracefully. Nevertheless, the work still supports the conclusion of our results that the IDS model which omits data flow is vulnerable to mimicry attacks.

3.2 Automated Mimicry Attack Construction

3.2.1 Definitions

Before explaining our mimicry attack generation algorithm, let us first establish some definitions. A *trace* is a sequence of system calls invoked by a program during its execution. In our context, a trace can be viewed simply as a string over some defined alphabet. In the Self-based IDS model [60, 140, 141], the alphabet for traces are the system call numbers. A *normal trace* is the trace generated by a program during its training stage, which represents the normal behavior of the program. This normal trace is used to construct the *normal profile* of the program.

The objective of a Self-based IDS is to examine traces and determine whether they are normal or anomalous based on the program's normal profile. We will call the attack trace which is detected by the IDS a *basic attack trace*. A mimicry attack disguises a basic attack trace into a *stealthy attack trace* which the IDS classifies as being normal.

In our analysis, we will also look at *subtraces*, which are simply substrings of a trace. In addition, we also consider *subsequences*, which are a subset of letters from a trace that is arranged in the original relative order. Thus, subsequences are different from subtraces, i.e. the former need not be contiguous in the trace.

3.2.2 Pseudo Subtraces

A weakness of a Self-based IDS which makes use of a normal profile represented as a set of k-grams is that it can accept subtraces which actually do not occur in the normal trace(s). For example, consider the following two subtraces of a normal trace:[2]

$$\langle \ldots, m_{i-4}, m_{i-3}, m_{i-2}, m_{i-1}, \mathtt{A}, \mathtt{B}, \mathtt{C}, \mathtt{D}, \mathtt{E}, \ldots$$
$$\ldots, \mathtt{B}, \mathtt{C}, \mathtt{D}, \mathtt{E}, \mathtt{F}, n_{i+1}, n_{i+2}, n_{i+3}, n_{i+4}, \ldots \rangle$$

Suppose the window size is 5, and assume that the subtrace $\langle \mathtt{A}, \mathtt{B}, \mathtt{C}, \mathtt{D}, \mathtt{E}, \mathtt{F} \rangle$ never occurs in the normal trace. This subtrace, however, will be accepted as normal by a Self-based

[2] For easier illustration, we opt to use English letters, instead of numbers, for representing system calls in the example trace(s).

IDS since the two 5-grams derived are *present* in the normal profile. We call such a subtrace *a pseudo subtrace*, which is defined as follows.

Definiton 3.1 (Pseudo subtrace). A subtrace is a *pseudo subtrace* for sliding-window size k if it is not a substring of the actual normal trace, yet passes the IDS detection since all of its k-grams are present in the normal profile.

A pseudo subtrace can be constructed by finding a common substring of length $k-1+\ell$ with $\ell \geq 0$ in two separate subtraces of length m ($\geq k+\ell$) and n ($\geq k+\ell$) respectively, and then joining them to form a new subtrace of length $m+n-k-\ell+1$. For our attack construction, we set $\ell = 0$ so as to put the weakest constraint on the mimicry attack construction with sliding-window size k.

Example 3.1. Given two unrelated subtraces $\langle \mathtt{A}, \mathtt{B}, \mathtt{C}, \mathtt{D}, \mathtt{E} \rangle$ and $\langle \mathtt{B}, \mathtt{C}, \mathtt{D}, \mathtt{E}, \mathtt{J}, \mathtt{K} \rangle$ with $k = 5$, we can construct a pseudo subtrace $\langle \mathtt{A}, \mathtt{B}, \mathtt{C}, \mathtt{D}, \mathtt{E}, \mathtt{J}, \mathtt{K} \rangle$ as illustrated in Figure 3.1.

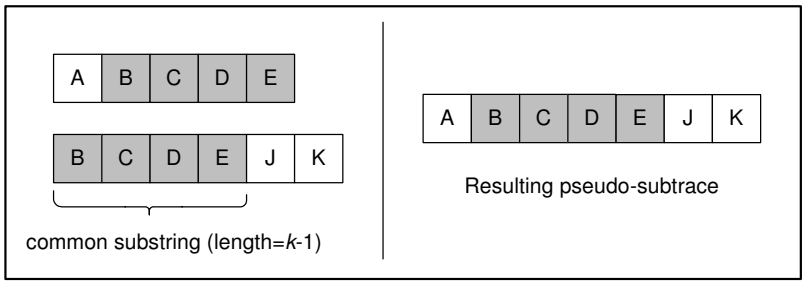


Figure 3.1: An example of pseudo subtrace construction with $k = 5$.

We can concatenate a pseudo subtrace with a normal subtrace, or another pseudo subtrace, to create a longer pseudo subtrace. *A stealthy attack trace* version of a basic attack trace is simply a pseudo subtrace in which the basic attack trace is its subsequence. Figure 3.2 illustrates such a complete process which combines subtraces into a stealthy attack trace containing basic attack sequence interspersed with no-ops.

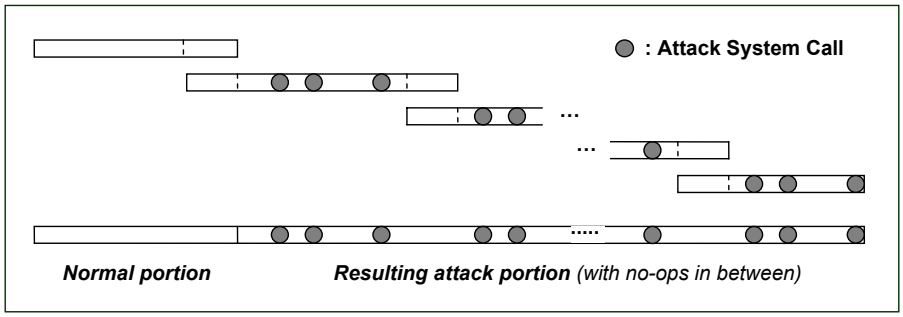


Figure 3.2: Mimicry attack construction by composing pseudo subtraces.

In our work, we use the term *pseudo subtrace* to specifically refer to the resulting subtrace which is obtained by joining two separate subtraces according to Definition 3.1. This resulting subtrace contains a *foreign sequence* (of *foreign order* type in the terminology of [154, 153]) of length $k + 1 + \ell$ as a substring. When $\ell = 0$ in the joining operation, the foreign sequence is a *minimal foreign sequence*. In the previous example, $\langle \mathtt{A}, \mathtt{B}, \mathtt{C}, \mathtt{D}, \mathtt{E}, \mathtt{F} \rangle$ is a minimal foreign sequence of length 6 for $k = 5$. The process in Figure 3.2 constructs a pseudo subtrace for a mimicry attack where multiple minimal foreign sequences of length $k + 1$ may exist along that subtrace. Each minimal foreign sequence combines two unconnected subtraces of normal traces together. Here, we emphasize the fact that the core components of mimicry attacks depend on the notions of subtraces and subsequences.[3]

3.2.3 The Overlapping Graph Representation

Given a normal trace, we represent the corresponding normal profile using a construct called an *overlapping graph*. This is similar to the *De-Bruijn* graph, which has been applied to various problems such as the "sequencing by hybridization" problem in computational biology [124].

Consider the normal trace of a program P, which is denoted as $N : \langle N_1, N_2, N_3, ..., N_n \rangle$, where N_i $(1 \leq i \leq n)$ is the letter representing a system call. We first augment the trace N by adding a suffix consisting of the $k - 1$ occurrences of sentinel symbol ('$'), signifying the end of the trace. This adds some additional k-grams, and is done to simplify the construction algorithm. Now, let K be the set of all k-grams derived from trace N according to the profile generation rule of the Self-based IDS. Given two strings p and q, the function $overlap(p, q)$ gives the maximal length of a suffix of p that matches a prefix of q. We define the *overlapping graph* constructed from K as follows.

Definiton 3.2 (Overlapping Graph). The *overlapping graph* G for a normal profile K which is generated with sliding-windows size k is defined as: a directed graph (V, E), where the vertices V are the k-grams in K, and $E \subseteq V \times V$ with each edge $e \in E$ connects two vertices p and q whenever $overlap(p, q) = k - 1$.

Example 3.2. Given a normal trace $N : \langle \mathtt{A}, \mathtt{B}, \mathtt{C}, \mathtt{D}, \mathtt{E}, \mathtt{F}, \mathtt{G}, \mathtt{A}, \mathtt{B}, \mathtt{E}, \mathtt{F}, \mathtt{H} \rangle$ with k=3, the corresponding overlapping graph G is as shown in Figure 3.3. For simplicity, we do not show the nodes corresponding to 3-grams of $(\mathtt{F}, \mathtt{H}, \$)$ and $(\mathtt{H}, \$, \$)$ which are in G.

Notice that there are two types of edges in G: *direct edges* and *pseudo edges*. The direct edges are those edges which connect vertices p and q corresponding to two *con-*

[3]We remark that not all foreign sequences according to the definition in [154, 153] are related to the notion of pseudo subtraces. We are not concerned with foreign sequences that contain system calls not in the k-grams. This is because they cannot be used to generate mimicry attacks.

secutive k-grams derived from the normal trace. Hence, both p and q contain a common substring of length $k-1$ which stems from *the same* substring in the normal trace. Pseudo edges are those which are not created due to the same substring of length $k-1$ in the normal trace. A pseudo edge is of special importance as it can be used to generate a pseudo subtrace. In Figure 3.3, the direct edges are drawn with single arrows, while the pseudo edges are drawn with double arrows.

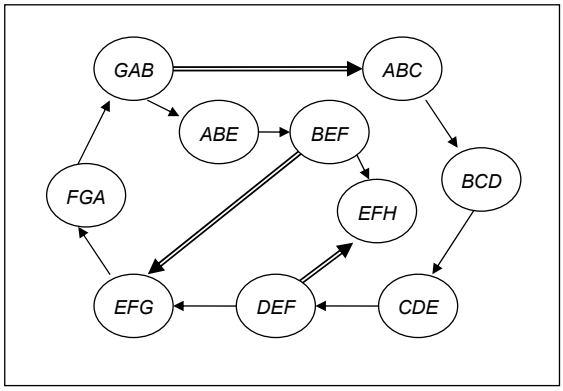


Figure 3.3: Overlapping graph G for $N : \langle \mathtt{A}, \mathtt{B}, \mathtt{C}, \mathtt{D}, \mathtt{E}, \mathtt{F}, \mathtt{G}, \mathtt{A}, \mathtt{B}, \mathtt{E}, \mathtt{F}, \mathtt{H} \rangle$ with $k = 3$. For simplicity, nodes corresponding to 3-gram $(\mathtt{F}, \mathtt{H}, \$)$ and $(\mathtt{H}, \$, \$)$ are not shown.

The graph G can also be viewed as an FSA model for recognizing normal traces. The work [167] develops an FSA from the k-gram database profile. Their approach however does not distinguish between what in the overlapping graph corresponds to direct and pseudo edges. Since our concern is to address the limitations of the Self-based IDS, the overlapping graph gives us a natural capability where we can evaluate the difference between allowing pseudo edges and removing them.

So far, by our definition of G, we always assume that all system calls can be turned into no-ops. Wagner and Soto [167] analyze system calls in Unix/Linux environment, and conclude that apart from just a few system calls, almost all can be easily nullifiable. The exception applies mainly to `exit()`, `pause()` and `vhangup()`. We thus can define S^- as the set of unnullifiable system calls. To avoid any subsequent effect of using S^- in the stealthy trace, we can remove all outgoing edges from every node whose k-gram starts with any letter in set S^-. Thus, a stealthy trace can only contain a system call $s \in S^-$ as its last system call in the trace.[4]

[4]If deemed necessary, we can additionally remove all incoming edges into every node whose k-gram starts with any letter in set $S^- - \{$ `exit()`$\}$ to avoid executing unnullifiable system calls altogether.

3.2.4 Mimicry Attack Construction

Now, we would like to construct mimicry attacks by using the overlapping graph. Given an overlapping graph G and a basic attack trace $A : \langle A_1, A_2, A_3, \ldots, A_\ell \rangle$ which is detectable by the IDS, we want to automatically construct the shortest stealthy attack trace $L_{min} : \langle L_1, L_2, L_3, \ldots, L_m \rangle$, where $m \geq \ell$, which contains $A_1, A_2, A_3, \ldots, A_\ell$ as a subsequence and where the other system calls in $\{L_{min} - A\}$ behave as no-ops with respect to A.

We consider automated attack constructions in two following attack scenarios:

1. *Trojan attack scenario*: where the attacker replaces a victim program with a trojan that executes a stealthy attack trace.
2. *Code-injection attack scenario*: where the resulting stealthy attack trace is to be injected into a victim program's process memory, and subsequently be executed. This can be achieved, for example, by using a buffer overflow attack technique, where the stealthy trace is crafted as the shellcode of the buffer-overflow exploit.

The code-injection attack scenario poses more challenges since the transition between the pre-injection subtrace and the stealthy attack trace must still pass the IDS detection. In addition, from the attacker's viewpoint, the shortest stealthy attack trace is desirable in this scenario for the following important reasons. Firstly, an attack trace with a minimum number of system calls will achieve attack efficiency. Secondly, and more importantly, a rather long shellcode injected as input data into the program may be noticed by other security measure(s) deployed on the host. A security measure that monitors the length of an input fed into the program may suspect such long shellcode as a malicious input, or an anomalous request [82]. Lastly, the most straightforward buffer overflow technique is one which puts the whole shellcode inside the overflowed buffer [119]. Hence, for this technique to work, the shellcode is restricted by size constraints.

3.2.5 Attack Construction Algorithm under Trojan Attack Scenario

In a trojan attack scenario, transforming a basic attack trace A into *a stealthy attack trace* L is equivalent to:

> Finding a path P on the overlapping graph G which monotonically visits nodes whose k-gram label begins with the symbol A_i for all $1 \leq i \leq \ell$.

The shortest stealthy trace L_{min} is a stealthy attack trace with the minimum number of system calls among all the stealthy attack traces.

To allow for an attack construction, we augment the previously constructed overlapping graph G into the *extended overlapping graph* G' defined as follows.

Definiton 3.3 (Extended Overlapping Graph). The *extended overlapping graph* $G' = (V + W, E + Occ)$ is the resulting graph constructed by augmenting G with an additional subgraph called the *occurrence subgraph* $O = (W, Occ)$. The nodes in the occurrence subgraph, W, are the unique individual letters from the k-grams in G. The edges Occ, which connect W and V, are constructed as follows: for each node $w \in W$, we add an outgoing edge to every node in G where the first letter in its k-gram label is the same as the letter for w.

Example 3.3. Figure 3.4 shows the extended overlapping graph for the overlapping graph in Example 3.2 (Figure 3.3).

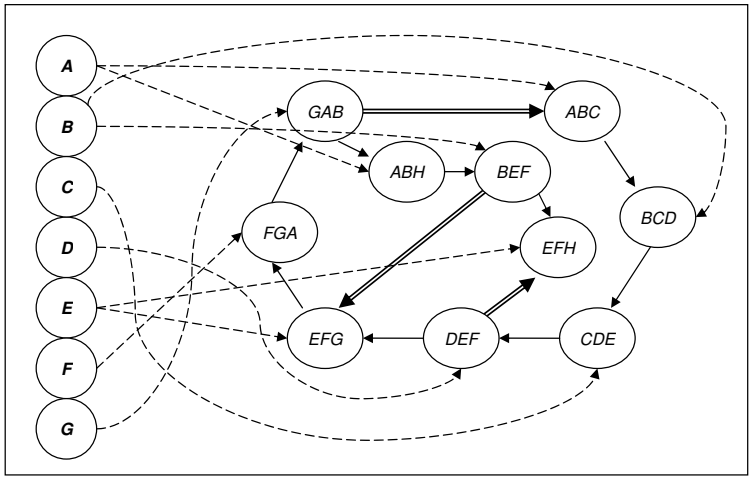


Figure 3.4: Extended overlapping graph G' from graph G in Figure 3.3.

Before outlining the construction algorithm, we illustrate the mimicry attack construction with the following example.

Example 3.4. Suppose that we want to construct a stealthy attack trace from a basic attack trace $A : \langle \texttt{G}, \texttt{C}, \texttt{D} \rangle$ using the extended overlapping graph G' in Figure 3.4. Note that trace A is detected by the IDS since its corresponding 3-gram $(\texttt{G}, \texttt{C}, \texttt{D})$ is not in the normal profile. Inspecting graph G', we however find the stealthy path: `GAB-ABC-BCD-CDE-DEF`. Thus, the stealthy attack trace is the sequence of $\langle \underline{\texttt{G}}, \texttt{A}, \texttt{B}, \underline{\texttt{C}}, \underline{\texttt{D}} \rangle$, with `A` and `B` added as no-ops. This example uses the pseudo edge (`GAB`, `ABC`).

Our attack construction builds a search tree, and performs a search on the tree to find the shortest mimicry attack. Except the root node, each node in the search tree corresponds to one letter in the basic attack trace, i.e. A_i with $1 \leq i \leq \ell$. The branches from A_i where $1 \leq i < \ell$ are the choices of extending a path on graph G corresponding to subtrace $\langle A_1, \ldots, A_i \rangle$ into one corresponding to $\langle A_1, \ldots, A_i, A_{i+1} \rangle$. A stealthy attack trace is found when we can reach A_ℓ which is the last required attack system call.

In order to make the search more efficient, we employ a *branch-and-bound* strategy to prune the constructed attacks which exceed the best solution found so far. Our implementation constructs an all-pair shortest-path table, which is used both to test connectivity between two nodes in G and also to assist in pruning for the branch-and-bound search. A sketch of the algorithm is given in Algorithm 3.1.

Algorithm 3.1 Attack construction on Self-based IDS under trojan attack scenario

Input:
- Extended overlapping graph G' constructed from the program's normal profile K
- A normal trace $N : \langle N_1, N_2, N_3, \ldots, N_n \rangle$
- Basic attack trace $A : \langle A_1, A_2, A_3, \ldots, A_\ell \rangle$

Output:
- Shortest stealthy attack trace $L_{min} : \langle L_1, L_2, L_3, \ldots, L_m \rangle$
- Or failure, if no solution trace can be found.

1. Construct the *all-pair shortest-path table*, with the following initialization:
   Between any two adjacent nodes p and q, set distance(p,q):=1.
   If two nodes p and q are not connected Then distance(p,q):= ∞.
2. $Min\_distance := \infty$ and $Min\_trace := \langle\rangle$.
   Create a special node v_0 as the root of the search tree, and $\forall\, v \in V$, $distance(v_0, v) := 0$.
3. (Perform *branch-and-bound* search on the search tree as follows:)
   current_cost:= 0 (it keeps track of the distance from v_0 to v_i in the path being explored)
   For all i:=1 to ℓ choose v_i from $\{v_i | (A_i, v_i) \in Occ\}$:
   - If $distance(v_{i-1}, v_i) = \infty$ Then backtrack.
   - Add $distance(v_{i-1}, v_i)$ to *current_cost*.
   - If *current_cost* $\geq$ *Min_distance* Then backtrack.
   - If complete solution is found ($i = \ell$) Then
     If *current_cost* $<$ *Min_distance* Then
     Min_distance := *current_cost*;
     Min_trace := *current_trace*.
4. Once the search tree is fully explored:
   If *Min_distance* $= \infty$ Then return failure;
   Else return *Min_trace* (as the shortest stealthy attack trace L_{min}).

3.2.6 Attack Construction Algorithm under Code-Injection Attack Scenario

Under the code-injection attack scenario, a stealthy attack trace must be introduced only at the "*attack-introduction point*", that is after the code injection succeeds and the victim's program control flow is gained. Here, we need to make note of the k system calls prior to the attack point, which we call the *border sequence* $B : \langle b_1, b_2, \ldots, b_{k-1}, b_k \rangle$. This sequence corresponds to a *border k-gram* in the normal profile; which then translates into a particular node v_s in G, which we call a *border-start node*.

We then need to ensure that the concatenation of the pre-injection trace and the stealthy attack trace still passes the IDS. To this end, we extend Algorithm 3.1 as follows:

1. The border sequence $B : \langle b_1, b_2, \ldots, b_{k-1}, b_k \rangle$ must be included as an input to the search algorithm.
2. Given the border-start node v_s and border sequence B, we determine the *border-end node* v_z as follows:
   Find a path $J : j_1, j_2, \ldots, j_{k-1}, j_k$ (of length k-1) on G, where $j_1 = v_s$ and for $2 \leq i \leq k$, j_i is a node whose label starts with b_i.
   The last node in the resulting path, i.e. j_k, is the border-end node v_z.
3. The first-level nodes in the search tree, i.e. set $\{v \,|\, (A_1, v) \in Occ\}$, are explored during the search only if they are connected to the border-end node v_z.

3.2.7 Proof of Optimality of the Attack Construction

The construction algorithm produces the *shortest* stealthy attack trace for a given set of k-grams as the normal profile database, with k being the sliding-window size. Note that there may be more than one shortest stealthy attack trace, all of the same length. Our attack algorithm returns one of them.

We now prove the optimality of the attack construction algorithm for both the trojan and the code-injection attack scenarios.

Theorem 3.1 (Stealthy Attack Trace Optimality). Let k be the chosen sliding-window size and K be a set of k-grams as the normal profile database, then the stealthy attack trace $L_{min} : \langle L_1, L_2, L_3, ..., L_m \rangle$ returned by the attack construction algorithm is the *shortest* stealthy attack trace of the basic attack trace $A : \langle A_1, A_2, A_3, \ldots, A_\ell \rangle$.

Proof. We will prove the optimality by contradiction. Suppose there exists as a stealthy attack trace that is shorter than L_{min}, which is denoted by $L' : \langle L'_1, L'_2, L'_3, ..., L'_n \rangle$, with $n < m$. Given that all k-grams derived from trace L' must be present in the set K, there exists a node corresponding to each k-gram of L' in the overlapping graph G, which is constructed based on K.[5] Moreover, since any two consecutively generated k-grams share a substring of length $k-1$, then the two nodes corresponding to these two consecutive k-grams are connected by an edge in G. As a result, the stealthy trace L' corresponds to a path $P' : v'_1, \pi'_1, v'_2, \pi'_2, v'_3, \pi'_3, \ldots, v'_{\ell-1}, \pi'_{\ell-1}, v'_\ell$ on G, where the following conditions hold (here v'_i, $1 \leq i \leq \ell$, is a node in G; and π'_i, $1 \leq i < \ell$, denotes a path on G with zero or more nodes):

1. For each node v'_i on path P', with $1 \leq i \leq \ell$, the first letter of its label is the attack system call A_i.

[5] Recall that graph G also contains $k-1$ nodes whose label contains sentinel symbol ('\$') due to our addition of $k-1$ occurrences of sentinel at the end of the trace (see Section 3.2.3).

2. For each node v'_i, $1 \leq i \leq \ell$, there exist a node $w_i \in W$ whose label is A_i and an edge $(w_i, v'_i) \in Occ$ connecting w_i to v'_i.
3. Node v'_i, $1 \leq i < \ell$, is connected to node v'_{i+1} through a path π'_i on G, which consists of zero or more nodes.
4. Additionally, under the code-injection attack scenario (Section 3.2.6), there exists a path from the border-end node v_z to node v'_1.

Following the same reasoning on the correspondence between a stealthy attack trace and a path on G, we also have a path P_{min} on G which corresponds to the reported stealthy trace $L_{min} : \langle L_1, L_2, L_3, \ldots, L_m \rangle$. Due to the 1-to-1 correspondence between a system call in a stealthy attack trace and a node in the corresponding path on G, the number of system calls of a stealthy trace is the same as the number of nodes on the corresponding path. Here, we use $|P|$ to denote the number of nodes in a path P. Note that $|P|$ = 1 + the length of path P. We thus have $|P'| = n$, $|P_{min}| = m$, and $n < m$.

By the definition of our search tree in the attack construction algorithm, there exists a path P_{equiv} in the search tree which also contains $v'_1, v'_2, v'_3, \ldots, v'_{\ell-1}, v'_\ell$ as its subsequence. That is, P_{equiv} is in the form of: v'_1, π_1, v'_2, π_2, v'_3, $\pi_3, \ldots,$ $v'_{\ell-1}$, $\pi_{\ell-1}$, v'_ℓ; with π_i, $1 \leq i < \ell$, denoting a path on G with zero or more nodes.

Now let us consider paths P' and P_{equiv}. We have two possible cases:

1. If $|P_{equiv}| \leq |P'|$:
   Since the branch-and-bound algorithm returns a trace with the shortest length among all solution paths, we have $|P_{min}| \leq |P_{equiv}|$. By transitivity, we have $|P_{min}| \leq |P'|$. Thus, $m \leq n$. However, we assume earlier that L' is shorter than L_{min}, that is: $n < m$. This leads to a contradiction.

2. If $|P_{equiv}| > |P'|$:
   Since both paths share a common subsequence: $v'_1, v'_2, v'_3, \ldots, v'_{\ell-1}, v'_\ell$, the path P' can be shorter than P_{equiv} only if:

$$\sum_{i=1}^{\ell-1} |\pi'_i| < \sum_{i=1}^{\ell-1} |\pi_i| \tag{3.1}$$

   However, the attack construction algorithm always makes use of the shortest path between any two nodes v'_i to v'_{i+1}, for $1 \leq i \leq \ell - 1$, to produce P_{equiv}. Hence, we have $|\pi_i| \leq |\pi'_i|$ for all $1 \leq i \leq \ell - 1$. As a result,

$$\sum_{i=1}^{\ell-1} |\pi_i| \leq \sum_{i=1}^{\ell-1} |\pi'_i| \tag{3.2}$$

   Therefore, $|P_{equiv}| \leq |P'|$. This contradicts our case assumption that $|P_{equiv}| > |P'|$.

Hence, the stealthy attack trace L' which is shorter than L_{min} does not exist. □

3.3 IDS Attack Experiments

3.3.1 Experimental Set-Up

We perform an automated stealthy attack construction under the trojan and the code-injection attack scenario on the following two variations of the Self-based IDS models:

- **Set Represented Self-based IDS (SET-IDS)**: where the normal profile is represented as a *set* of k-grams. This is the Stide model used in [60, 169].
- **Graph Represented Self-based IDS (GRA-IDS)**: where the normal profile is a *graph* of k-grams. With respect to the overlapping graph defined in Definition 3.2, now only the direct edges are allowed in this graph profile.

Our attack construction is implemented in C on a modest PC with a Pentium IV processor (1.82 GHz) with 256 MB of RAM running Red Hat Linux. We have also used various older versions of the Red Hat Linux distribution in order to obtain the traces of sample vulnerable programs together with their exploits. The traces are captured in Linux by using the `strace` utility. For simplicity, we have omitted system calls related to signal events, such as SIGALRM and SIGCHLD, due to their asynchronous nature.

The following remarks apply to our experiments:

- To construct a mimicry attack, the attacker requires the knowledge of the vulnerable program's actual profile (K) in the victim host. Generally, K is not available to the attacker. Yet, given possible information and insights on the OS environments and the program's configurations, the attacker can attempt to simulate the targeted program and derive an approximated profile (K') to be as close as possible to K. In our experiments, we use K as an input to the algorithm so as to investigate the actual susceptibility of the programs against mimicry attacks. This assumption is relevant particularly when the smart attacker assumed could derive a good approximation of K, or when it somehow manages to obtain K.

- The three exploits below make use of `execve()` system call to spawn a root shell. However, `execve()` is not present in the normal trace. Therefore, we use an alternative strategy to write an entry into the file "`/etc/shadow`". This actually corresponds to *Attack-strategy* A_2 from our list of strategies given in Table 4.2 (see Section 4.3.2). This particular attack strategy is chosen for comparison here as it has been used for mimicry attacks on the Self-based IDS before (e.g. [152]). We remark that it is perfectly reasonable to modify the original attack since we assume an intelligent adversary.

3.3.2 Sample Vulnerable Programs and Attack Construction

As our goal is to investigate the practicality of our automated attack construction, we experiment with real programs using real available exploits.

Traceroot2 (Traceroute Exploit)

This traceroute exploit was previously used in [152]. It is available at `http://www.packetstormsecurity.org/0011-exploits/traceroot2.c`. The exploit attacks LBNL traceroute v1.4a5 which is included in the Linux Red Hat 6.2 distribution.

The original attack sequence is: `setuid(0),setgid(0), execve("/bin/sh")`. This is changed into: `open(),write(), close(),_exit()` as in [152]. The results of the attack construction on normal traces generated from three traceroute's sessions (with a total of 2,789 system calls) for window sizes from k=5 to 11 are given in Table 3.1.

Traceroute Result	**k=5**	**k=6**	**k=7**	**k=8**	**k=9**	**k=10**	**k=11**
Length of Resulting Stealthy Attack Trace:							
SET-IDS (Code-Injection)	46	48	48	64	64	112	116
GRA-IDS (Code-Injection)	48	48	64	64	116	116	125
SET-IDS (Trojan-Attack)	41	45	45	51	51	54	54
GRA-IDS (Trojan-Attack)	43	45	51	51	54	54	56
Average Search Time (User+Sys)	0.170s	0.210s	0.250s	0.300s	0.460s	0.388s	0.340s

Table 3.1: Attack construction results for traceroute with k=5 to 11 (with 2,789 system calls in the normal trace). SET-IDS and GRA-IDS represent the Self-based IDSs with the normal profile stored as a *set* of k-grams and a *graph* of k-grams respectively.

As can be seen in Table 3.1, the attack construction algorithm on traceroute is able to find stealthy attack traces on the two IDS variants and under the two attack scenarios. The shortest stealthy attack traces under the trojan attack mode are of below 60 system calls even with k=11. The stealthy attack traces under code-injection attack scenario require longer traces than those under trojan attack scenario given the same k. Storing the normal profile as a graph of k-grams (GRA-IDS) does not make the Self-based IDS substantially more robust against mimicry attacks. It makes the resulting stealthy traces only slightly longer than those required in the SET-IDS.

JOE Text Editor Exploit

The victim program that we chose is a popular Linux terminal text editor JOE, available at `http://sourceforge.net/projects/joe-editor/`. The exploit for Red Hat Linux was available at `http://www.uhagr.org/src/kwazy/UHAGr-Joe.pl`[6], and was run on Red Hat 7.3. JOE is not normally run as a setuid program. As a proof of

[6]The site seems to be no longer in operation (as of March 2010).

concept, we assume that JOE has been run from root (or setuid to root). The original attack sequence is: `setuid(0), execve ("/bin/sh")`. Again, we changed it to: `open(), write(), close(), _exit()`.

The results of the attack construction algorithm on JOE's normal traces generated from three JOE sessions (with a total of 9,802 system calls) for sliding-window sizes from k=5 to 11 are given in Table 3.2.[7]

JOE Result	k=**5**	k=**6**	k=**7**	k=**8**	k=**9**	k=**10**	k=**11**
Length of Resulting Stealthy Attack Trace:							
SET-IDS (Code-Injection)	20	30	49	76	79	80	81
GRA-IDS (Code-Injection)	30	49	76	79	80	81	82
SET-IDS (Trojan-Attack)	7	7	7	7	7	7	7
GRA-IDS (Trojan-Attack)	7	7	7	7	7	7	7
Average Search Time (User+Sys)	0.258s	0.305s	0.362s	0.432s	0.520s	0.623s	0.778s

Table 3.2: Attack construction results for JOE with k=5 to 11 (with 9,802 system calls in the normal trace).

Since JOE is a text editor, it falls into the class of general purpose programs as opposed to the more privileged processes targeted for monitoring by the Self-based IDS [60, 140]. We include it here to highlight some points on the results of attack construction. From the results shown in Table 3.2, we have made some interesting observations. Here, we find that for the trojan attack scenario, a stealthy attack trace of length 7 is sufficient for k=5 to 11. This is because a subtrace in the normal trace of JOE happens to contain the basic attack trace as its subsequence. Hence, *no* pseudo subtrace construction was done. However, this stealthy trace does not work under the code-injection attack scenario. Instead, longer stealthy attack traces are required.

Autowux WU-FTPD Exploit

This is the same exploit used in [167]. The `autowux.c` exploits "site exec" vulnerability on the WU-FTPD FTP server. It is available at `http://www.securityfocus.com/bid/1387/exploit/`. We ran the `wu-2.4.2-academ [BETA-15]` version of WU-FTPD that comes with Red Hat 5.0 distribution on the 2.2.19 kernel.

We use the same attack trace as [167] which is: `setreuid(),chroot(),chdir(), chroot(),open(),write(),close(),_exit()`. The results of the attack construction on the WU-FTPD normal traces generated from 10 sessions (11,051 system calls) for sliding-window sizes from k=5 to 11 are given in Table 3.3.

[7] From the normal traces collected for JOE, we note that there are actually some differences between the normal traces and the exploit trace before the attack-introduction point due to some `brk()` system calls. This is probably due to an increased memory allocation for the buffer overflow attack. However, as reasoned by [167], small differences may be tolerated by the Self-based IDS depending on the parameters used in its anomaly-signal measurement function (e.g. Locality Frame Count).

WU-FTPD Search	k=**5**	k=**6**	k=**7**	k=**8**	k=**9**	k=**10**	k=**11**
Length of Resulting Stealthy Attack Trace:							
SET-IDS (Code-Injection)	92	182	196	230	256	272	321
GRA-IDS (Code-Injection)	182	194	212	244	272	303	328
SET-IDS (Trojan-Attack)	77	167	181	201	234	257	285
GRA-IDS (Trojan-Attack)	167	179	183	222	257	285	314
Average Search Time (User+Sys)	2.036s	2.663s	3.535s	5.056s	4.980s	6.220s	7.811s

Table 3.3: Attack construction results for WU-FTPD with k=5 to 11 (with 11,051 system calls in the normal trace).

We can see from Table 3.3 that the resulting stealthy attack traces happen to be relatively longer than those in the first two sample programs. The work [167] give a stealthy trace for k=6 with 135 system calls based on their normal profile. Their result, however, is not comparable to ours as the normal traces used are different. In their case, they had collected normal traces for an existing WU-FTPD with large numbers of downloads over two days. On the other hand, we have used a small normal profile.

We additionally discuss some patterns which commonly apply to the three sample programs below.

3.4 IDS Evaluation Discussion

We have shown that the two variants of the Self-based IDS are vulnerable against mimicry attacks even with sliding-window size longer than that is usually employed (k=6 to 8). The running times also show that our automated attack construction algorithm is practical and efficient. Execution times for all cases is at most a few seconds even on relatively large window sizes.

Based on the results, we also observe the following common trends:

- There can be a considerable difference in length between the stealthy code-injection attack traces and the trojan attack ones. In some cases, like in JOE, the stealthy attack trace under the trojan attack scenario is very short. Here an attack of length 7 works for window sizes from k=5 to 11.[8]

- The length of the shortest stealthy attack trace varies from program to program. It confirms the earlier reports [152, 50] that a larger window size (k) tends to require also a longer stealthy attack trace. In practice, however, there is a limit on the choice of k due to the increase in normal profile size and processing overheads of the IDS [169, 67]. Nevertheless, it clearly shows that relying on the Self-based IDS

[8]The actual trojans will usually have longer sequences since there are additional system calls invoked at the beginning of a program related to libraries loading or memory allocation. However, the reported number of system calls does establish the lower-bound of mimicry attacks in the trojan attack setting.

with a certain size of sliding-window of, such as 6 as suggested in [60], is insufficient. Rather, other improvements are necessary.

- We can see that removing pseudo edges on graph G for the GRA-IDS model does not make the Self-based IDS significantly stronger against mimicry attacks. In other words, pseudo subtraces can still exist even if we store the normal profile as a *graph* of k-grams. We find that, besides pseudo edges, there is another source of imprecision in the GRA-IDS model which would allow for a pseudo subtrace construction. To understand this, let us now consider a normal trace $\langle \mathtt{A}, \mathtt{B}, \mathtt{C}, \mathtt{D}, \mathtt{E}, \mathtt{A}, \mathtt{B}, \mathtt{C}, \mathtt{M} \rangle$ with $k = 3$. Figure 3.5 shows the graph without pseudo edges as the normal profile for this trace.

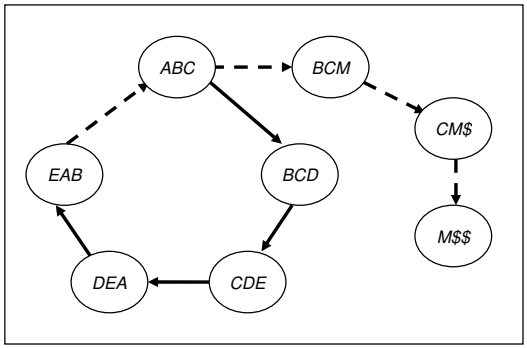


Figure 3.5: A graph of 3-grams (without pseudo edges) used in GRA-IDS model for the sample trace $\langle \mathtt{A}, \mathtt{B}, \mathtt{C}, \mathtt{D}, \mathtt{E}, \mathtt{A}, \mathtt{B}, \mathtt{C}, \mathtt{M} \rangle$. Note that a "common node" `ABC` allows for pseudo subtrace construction.

A stealthy attack trace $\langle \underline{\mathtt{E}}, \mathtt{A}, \underline{\mathtt{B}}, \mathtt{C}, \underline{\mathtt{D}} \rangle$ can be constructed for a basic attack trace $\langle \mathtt{E}, \mathtt{B}, \mathtt{D} \rangle$. This is achieved by a pseudo subtrace construction of $\langle \mathtt{E}, \mathtt{A}, \mathtt{B}, \mathtt{C}, \mathtt{D} \rangle$ from two unrelated subtraces $\langle \mathtt{E}, \mathtt{A}, \mathtt{B} \rangle$ and $\langle \mathtt{A}, \mathtt{B}, \mathtt{C}, \mathtt{D} \rangle$. In Figure 3.5, the constructed stealthy trace corresponds to the path: `EAB-ABC-BCD-CDE-DEA`. The reason why the path exist in the graph is that a "*common node*" `ABC` facilitates such a path construction. To illustrate this construction, Figure 3.5 is drawn with two different outgoing edges from node `ABC` based on the actual overlapping relationships with its adjacent nodes (k-grams) as they appear in the normal trace. A stealthy path can make use of a common node like `ABC` to combine two unrelated paths (shown as solid and dotted lines in the figure). Such common nodes exist due to the existence of a common substring of length k, e.g. $\langle \mathtt{A}, \mathtt{B}, \mathtt{C} \rangle$ in the example, between any two unrelated subtraces in the normal trace.

3.5 Using Attack Construction to Measure IDS Security

Although our main contributions focus on the automated attack construction on the Self-based IDS, an important point of our work is to demonstrate that the attack construction above provides a systematic method to effectively measure the *resistance level* of an IDS against targeted attacks. We describe our developed framework in this section.

3.5.1 Approach and General Framework

We develop our framework upon the notion of "*attack-space search*". An attack space is a state representation derived from the properties and detection mechanism of the IDS given that one of its security assumptions is broken. We will show how actual attack spaces look like by means of examples in the sections to follow. Once the attack search objective has been defined, the problem is then recast into how we can perform an efficient search on the attack space.

Our framework for evaluating the IDS security involves the following steps:

Step 1. Study and formalize the IDS' *definition and its security assumptions*;

Step 2. Find a *security assumption which can be made invalid*;

Step 3. Define the *attack space*;

Step 4. Establish the *search objective*;

Step 5. Construct a corresponding *search algorithm*;

Step 6. Apply the search on the *gathered experimental data*;

Step 7. (Optionally) modify the IDS model so that the search process becomes *computationally harder* or its search objective becomes *unfeasible* in the new attack space. This step thus increases the strength of the IDS against attacks which attempt to exploit the invalidated assumption.

Steps 1–5 above realize the objective of obtaining the attack crafting time as well as the constraints for a successful attack, which will be our main focuses here. Step 6 is the step which analyzes the attack occurrence given a normal profile dataset. Defining an attack space (Step 3) is an important step, which unfortunately requires some insights to be carried out. In defining an attack space, we can always come up with different representations for it. The efficiency of an attack construction, of course, depends on the right representation and its corresponding search algorithm.

Rather than giving detailed definition of the above steps, we instead show below how we can apply the framework. We revisit the attack construction on the Self-based IDS (Stide) by describing the construction process within the developed framework. Details of the attack process are now fitted into the framework's defined steps. This makes the higher-level strategy of constructing mimicry attacks on the Self-based IDS become much

clearer. Additionally, we then show how the framework can be applied to another IDS model.

3.5.2 Applying the Framework to Self-based IDS

Security Assumption

Self-based IDS (Stide) assumes that k-grams of a running program are good predictors of the program's behavior. More specifically, it assumes that:

> "It is *difficult* for an attacker to inject system call(s) into the program's system call trace without introducing *foreign k-grams* [9] which will be detected by the IDS."

Invalidating the Assumption

Stide's assumption above, however, can be made invalid by the following observation.

> "It is *quite possible* to combine two unrelated subtraces to make a longer subtrace that derives no foreign k-grams. That is, all of the k-grams of the combined subtrace are present in the program's normal database".

We have shown this constructively using our pseudo subtrace construction described in Section 3.2.2.

Attack Space, Search Objective, and Algorithm

Here, we define the attack space for Stide that fulfills the requirements that:

1. Two k-grams derived from $(k+1)$-length of the normal subtrace are related since they can be joined to produce (part of) a stealthy attack trace;
2. A pseudo subtrace can be constructed from two unrelated k-grams provided that they share a common substring of length $k-1$.

We have used overlapping graph for the attack space (see Section 3.2.3), in which the requirement 1 is satisfied by the direct edges, whereas requirement 2 is realized by the pseudo edges in the overlapping graph. Accordingly, we also have shown the search objective on attack space and the construction algorithm in Section 3.2.

3.5.3 Applying the Framework to the FSA-based IDS

We now examine a different IDS model, which is based on FSA as proposed by Sekar et al. [136]. We have briefly summarized this IDS model in Section 2.2.3. In our attack construction, we make use of the technique to forge a Program Counter (PC) through

[9] Foreign k-grams are k-grams which are not present in the normal profile database.

a buffer overflow attack as suggested in [50]. Hence, we extend the results here by formalizing the search space for the IDS, and constructing the search algorithm using our framework.

Security Assumption

The FSA-based IDS incorporates the PC information into the IDS model. It assumes that the PC information can be reliably obtained, and its inclusion can enhance the detection and security of the IDS. More specifically, we can state this security assumption as follows:

> "PC location represents a program point where a system call is made within the program. It can be accurately obtained, for example by using a stack traversal mechanism. The attacker *cannot* forge the stack (or other runtime environment information) in order for the stack traversal to recover the correct PC information."

Invalidating the Assumption

The work [50] points out how one can forge PC and return address through a buffer overflow attack.[10] The attack technique makes use of `ret` instead of `call` instruction to launch an exploit system call without inserting a stack frame into the stack. The hijacked control flow is then made to jump and re-execute the vulnerable operation[11] again and again until the complete exploit sequence has been executed. To achieve one system call execution, three forged stack frames are inserted into the stack in the following "push order": (i) a stack frame with a PC address of the vulnerable function; (ii) another stack frame with a PC that falls within the program's code region; and (iii) one with a PC of the exploit system call somewhere within the library. Note that since we assume the occurrence of buffer overflow in the user space, such a vulnerable operation does not involve any invocation of system call.

From our framework's viewpoint, the attack technique actually provides a mechanism to invalidate IDS' security assumption on the stack integrity and the correctness of the recovered PC addresses. Since the IDS does not check system call arguments, an attack will therefore successfully fool the IDS as long as its transitions (PCs and executed system calls) are accepted by the normal FSA.

[10] A technique proposed by Kruegel et al. [80] can additionally be used on the program's binary to automate the process of modifying code pointers following a system call so as to transfer the execution control back to the attacker for the invocation of subsequent attack system calls.

[11] We assume that the overflow occurs in the user space due to the program's use of an unsafe function.

Attack Space and Search Objective

Different from the Self-based IDS (Stide) where we need to define the overlapping graph, the FSA-based IDS generates an FSA as the normal representation, which directly provides the attack space for the IDS. We define an FSA representing the IDS normal profile as a tuple $(S, \Sigma, \delta, s_0, s_\perp)$[12], where: S as the set of states (PCs); Σ as a set of all system calls; s_0 as the initial state; $s_\perp$ as a special state indicating that an anomaly has occurred; and $\delta : S \times \Sigma \rightarrow S$.

The state and the system call before the buffer overflow occurs are of particular importance. Let us call that state s_{prev} and the system call c_{prev}. Given an FSA $(S, \Sigma, \delta, S_0, s_\perp)$, s_{prev}, c_{prev}, and an intended (basic) attack trace $A : \langle a_1, a_2, a_3, ..., a_\ell \rangle$, our search objective is thus to form the shortest path P on the FSA of the following form:

$$P_{min} = s_{prev} \xrightarrow{c_{prev}} s_i \xrightarrow{c_1} s_j \xrightarrow{c_2} s_k \ldots s_x \xrightarrow{c_{m-1}} s_y \xrightarrow{c_m} s_z \tag{3.3}$$

where: $m \geq \ell$; the stealthy attack trace $L = \langle c_1, c_2, c_3, ..., c_m \rangle$ contains A as a subsequence; and all system calls in $\{L - A\}$ behave as no-ops with respect to A. The terminating state s_z is the special state after the invocation of the last system call in the exploit (usually `_exit()`), thus allowing the exploit to make a grateful exit.

The key to the attack construction is the fact that we can forge the PC for a state executing an exploit system call so that it appears to fall within the valid program's code region, and that the forged PC is connected to the previous valid state in the FSA. The PC forging is achieved by the crafting of the second stack frame in the corrupted stack after a buffer overflow attack occurs as discussed in [50].

Search Algorithm

To facilitate the search, we add an auxiliary subgraph similar to the one added for the Self-based IDS (Stide) attack construction. *Occurrence subgraph* $O = (W, Occ)$ is defined with W ($\subseteq \Sigma$) as a set of unique system calls invoked in the FSA. For each node $w_i \in W$, we add an outgoing edge to all of its occurrences in δ; thus forming a set of edges Occ. Thus, unlike in Stide where a node $w \in W$ points to nodes, here it points to system call labels in the transitions.

Our attack construction algorithm performs the search starting from all the states that are (directly) adjacent to s_{prev} by c_{prev}. We denote this set as $B = \{b \in S \mid s_{prev} \xrightarrow{c_{prev}} b\}$. Hence, there may be multiple search trees, each with state $b \in B$ as the root node. For each b, the search incrementally expands the stealthy attack path formed so far to the next states s_i, where $\{s_i \mid s_{i-1} \rightsquigarrow \xrightarrow{c} s_i \quad \wedge \quad (a_i, c) \in Occ\}$, for $1 \leq i \leq \ell$. The starting state s_0 is set to b. Recall that a_i ($1 \leq i \leq \ell$) is a system call in the basic attack trace A.

[12]It is assumed that every state $s \neq q_\perp$ is an accepting state.

The connection symbol ('$\rightsquigarrow$') denotes a path of length zero or more in the FSA.

As before, we employ a branch-and-bound strategy to prune a constructed path which exceeds the best solution found so far. The shortest distance between any two states is stored in a "distance table". It is employed both to test connectivity between two states and also to assist in pruning for the branch-and-bound search. The algorithm for constructing attack on the FSA-based IDS is a variation of the one on the Self-based IDS (see Algorithm 3.1). They both make use of the same branch-and-bound search procedure. However, we use a different search tree definition to be explored in the algorithm for attacking the FSA-based IDS. A sketch of the algorithm to automatically construct a stealthy attack trace on the FSA-based IDS under code-injection scenario is shown in Algorithm 3.2.

Algorithm 3.2 Attack construction on the FSA-based IDS under code-injection scenario

Input:
- Normal-profile FSA (a directed graph);
- Basic attack trace $A : \langle a_1, a_2, a_3, ..., a_\ell \rangle$;
- The last state before the buffer overflow occurrence (s_{prev});
- The last system call before the first buffer overflow occurs (c_{prev}).

Output:
- Shortest stealthy attack path $P_{min} : s_{prev} \xrightarrow{c_{prev}} s_i \xrightarrow{c_1} s_j \xrightarrow{c_2} \ldots \xrightarrow{c_{m-1}} s_y \xrightarrow{c_m} s_z$;
- Or failure, if no solution trace can be found.

1. Construct the *all-pair shortest-path table*, with the following initialization:
   Between two adjacent states p and q, set distance(p,q):=1.
   If two states p and q are not connected Then distance(p,q):= ∞.
2. Set $Min\_distance := \infty$ and $Min\_path := \langle\rangle$.
3. For all $b \in \{s_{prev} \xrightarrow{c_{prev}} b\}$:
   (Perform branch-and-bound search on the search tree of b as follows:)
   Set s_o:=b.
   current_cost:= 0 (it keeps track of the distance from b to s_i in the path being explored)
   For all i:=1..ℓ choose s_i from $\{s_i \mid s_{adjacent\_i} \xrightarrow{c} s_i\}$ with $(a_i, c) \in Occ$:
   - If *distance*(s_{i-1}, s_i) = ∞ Then backtrack.
   - Add *distance*(s_{i-1}, s_i) to *current_cost*.
   - If *current_cost* $\geq$ *Min_distance* Then backtrack.
   - If complete solution is found ($i = \ell$) Then
     If *current_cost* $<$ *Min_distance* Then
     Min_distance := *current_cost*;
     Min_path := *current_path*.
4. Once the search tree is fully explored:
   If *Min_distance* = ∞ Then return failure;
   Else return *Min_path* (as the shortest stealthy attack path P_{min}).

Below, we give an illustrative example on the attack construction on the FSA-based IDS.

Example 3.5. Suppose that we have an FSA as shown in Figure 3.6, which is also used as an example in [136]. Suppose that the basic attack trace is $A : \langle a_5, a_1, a_4 \rangle$, with state 3 as the last state before the buffer overflow occurs, and a_1 is the corresponding system call invoked. Although the occurrence subgraph is not shown here, we can easily form the following attack path by using the search construction above.

$$4 \xrightarrow{a_2} 6 \xrightarrow{a_4} 8 \xrightarrow{\boxed{a_5}} 3 \xrightarrow{\boxed{a_1}} 5 \xrightarrow{a_3} 6 \xrightarrow{a_4} 8 \xrightarrow{a_5} 10 \xrightarrow{a_3} 11 \xrightarrow{\boxed{a_4}} 12.$$

The exploit system calls performing real attacks are shown inside boxes, whereas the rest are made as no-ops.

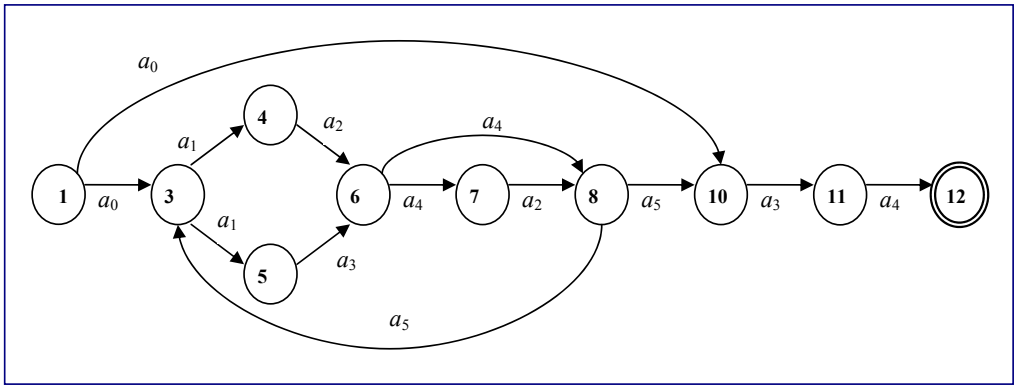


Figure 3.6: A sample FSA as a program's normal profile used in [136].

3.6 Summary

We have presented an efficient algorithm for automated mimicry attack construction on the Self-based IDS (Stide) and a variant model of it (GRA-IDS). Using several real programs and exploits, we also have shown the practicality of the attack construction, with execution times of at most a few seconds even for the relatively large window sizes. We also have shown how the construction method can be generalized into an approach to evaluate the robustness of an IDS against targeted attacks. The attack construction time is particularly useful to evaluate the resistance of an IDS against attacks in a zero-day attack setting. If we know that crafting an attack is possible but computationally hard, then we can have greater confidence that the IDS can function sufficiently well as the second line of defense during the program's unpatched time interval. This result is important since it shows how a quantitative measurement on IDS security strength against targeted attacks can be established.

Chapter 4

Improving Self-based IDS using Privilege and Argument Abstraction

The previous chapter has shown that the Self-based IDS is susceptible to mimicry attacks. In this chapter, we consider a simple enhancement to the Self-based IDS, which can either prevent mimicry attacks or make them more difficult. The enhancement makes use of system call arguments and process privilege information to improve the IDS. By using system call arguments, a data-flow aspect is incorporated into the IDS model. Taking the arguments into account is also important for mitigating non control-flow attacks, whose threat has been shown to be sufficiently realistic [35].

Besides being applicable to the Self-based IDS, our improved model can also strengthen various system-call monitoring IDS, such as the FSA-based [136] and the VtPath model [45]. To simplify our discussion, we will elaborate the enhancement to the Stide model of Self-based IDS, which keeps track of system call numbers in terms of a set of (full sequence) k-grams [60].

Our enhanced model has been published in [146], and also surveyed in [47]. The remainder of this chapter is organized as follows. Section 4.1 discusses the related works on improving the Self-based IDS using system call arguments. Section 4.2 elaborates our scheme which abstracts system call arguments and process privilege information. Our experiments are reported in Section 4.3, and the discussions on the experimental results are given in Section 4.4. Section 4.5 finally gives a summary of this chapter.

4.1 Related Works on Data-Flow based IDS

The general idea of analyzing arguments of operations for detecting behavior deviance appeared in a number of works. For example, [96] showed how the use of enriched

command-line data can enhance the detection of masqueraders. Our work reported in this chapter is based on the established Self-based IDS model (described in Section 2.2), and focuses on system call arguments. Below, we discuss several IDS models which also make use of system call arguments.

Kruegel et al. [83] made use of statistical analysis of system call arguments which can be used to evaluate features of the arguments such as string length, string character distribution, structural inference and token finder. It is however unclear whether the approach is sufficiently robust against targeted attacks such as mimicry attacks [17]. In addition, the work solely examined the arguments and disregarded the code flow altogether. The work was later extended in [112] by additionally using a Bayesian network to combine individual model scores into a single aggregate score. Another work which made use of machine learning technique on system call trace is [155]. It proposed a scheme to learn the attributes of system call arguments using a rule learning algorithm.

Taking a static analysis-based approach, Giffin et al. [52] made use of inter-procedural data-flow analysis to model statically-known arguments passed to system calls. Due to the nature of static analysis on program behavior, the scheme was only able to detect attacks that cause a program's runtime behavior to deviate from its statically extracted data-flow model. In our work, we focus on the gray-box and the black-box techniques which do not require the analysis of source codes or the binaries of the executables. Yet, the white-box technique and gray/black-box detector can complement each other as discussed in [87].

In a work proposed after ours, Bhatkar et al. [22] modeled the temporal aspects of data-flow by performing a learning on system call arguments. The learning establishes unary relations which correlate arguments with constant values, and binary relations of which each correlates two arguments of system calls on two different PC locations. Although our proposed technique does not relate arguments from two system calls, it has a particular strength in that the supplied simple policy represents an effective security model to prevent potentially dangerous operations according to their *direct effects* on the host's security. The work [22] did share the same conclusion as ours here that it is necessary to include data-flow model to be layered into the code-based IDS model.

Finally, there exist sandboxing techniques which also make use of system-call argument checking. The systrace scheme [125] uses system call policies to specify that certain system calls with specific arguments can be allowed or denied. This can be viewed as a the Self-based IDS with a window size of one, which is then enriched with system-call argument checking. We additionally remark that, compared to a basic policy definition used in [125], our policy definition (see Section 4.2.1) allows for a generalized mapping of arguments. Our scheme allows operations on files/directories to be grouped together into a set of specified categories according to their impact on the host's security.

4.2 Privilege and Argument Categorization (PAC) based IDS

In Unix, every process environment contains credentials which are evaluated by the OS access control mechanism when the process makes a system call. The credentials which determine the current privileges of a process are its effective user-id (euid) and effective group-id (egid). The euid/egid is either the actual real uid/gid of the user, or the one changed by invoking a setuid/setgid executable. Hence, euid and egid are simply a subset of all the user and group-id values defined in a system.

We propose to enhance k-grams to include not only the system number but also (abstracted) information about the euid, egid and system call arguments. We call our enhanced IDS model the *Privilege and Argument Categorization (PAC)-based IDS*. It is common for attacks to attempt exploiting programs while they are running in a privileged mode, or to elevate the privilege of the hijacked programs before performing damaging operations. The idea behind our privilege profiling is that such attacks can be detected if the corresponding executed system calls are unprivileged in the normal trace(s). This is particularly useful if a program conforms to a good setuid programming practice, which generally drops privileges as soon as they are no longer required. Rather than using the actual values, we abstract the euid, egid and system call arguments into categories based on a configuration specification. This is mainly to reduce the false positives which can be higher since the number of possible values is much greater. The abstraction technique also provides flexibility to group the arguments and the privileges together in terms of their importance/sensitivity levels.

4.2.1 Privilege and Argument Categorization

Formally, we can represent the privilege and argument categorization in the OS model with the following *mapping* functions:

- Function $EuidCat : U \rightarrow U'$, where: U = the set of euid and $U' \subset \mathbb{N}_0$ (the set of natural numbers starting from zero).
- Function $EgidCat : G \rightarrow G'$, where: G = the set of egid and $G' \subset \mathbb{N}_0$.
- Let $S \subset \mathbb{N}_0$ be the set of system call numbers in the OS model. For each $s \in S$, function $ArgCat_s : A_{s,1} \times A_{s,2} \times \ldots \times A_{s,no\_of\_args(s)} \rightarrow C_s$, where $A_{s,i}$ $(1 \leq i \leq no\_of\_args(s))$ = the set of possible entries for i-th argument of the system call s, and $C_s \subset \mathbb{N}_0$ as the categorized value for s and its arguments.

As mentioned earlier, we can formalize system-call monitoring IDS as an FSA [167, 50, 136]. Let us first define $s_z \in \mathbb{N}_0$ and $s_z \notin S$ to denote the "*sentinel*" system call. We

then define the extended system call set $S^+ = S \cup \{s_z\}$. The automaton based on our PAC-based IDS can be defined as $(\mathcal{Q}, \Sigma, \delta, q_0, \mathcal{Q} - \{q_\perp\})$, with:

- The set of states $\mathcal{Q} = \{q_0, q_\perp\} \cup \Sigma^k$, with k as the size of sliding window.
- The set of input $\Sigma = U' \times G' \times S^+ \times C$, with $C = \bigcup_{s \in S} C_s$.
- The initial state q_0.
- A special state $q_\perp$ which indicates that an anomaly has occurred. The other states $\mathcal{Q} - \{q_\perp\}$ are considered as accepting states.[1]
- The transition relation $\delta \subseteq \mathcal{Q} \times \Sigma \times \mathcal{Q}$ is constructed during the IDS training stage as follows: if the automaton is in state $q = (\sigma_1, \sigma_2, \ldots, \sigma_k)$, with $\sigma_i \in \Sigma \ \ (1 \leq i \leq k)$, and input $\sigma \in \Sigma$ is received, then (q, σ, q') is added to δ, with $q' = (\sigma_2, \ldots, \sigma_k, \sigma)$.

Here, we assume that a state $q \in \mathcal{Q} - \{q_0, q_\perp\}$ is always a k-tuple, i.e. $(\sigma_1, \ldots, \sigma_k)$. A complication however arises in the beginning or end of the normal trace, where only $i < k$ of σ_i entries are available. A workaround to this is by defining a "sentinel" entry $\sigma_z = (u', g', s_z, c)$, with s_z as the sentinel system call; and u', g', and c being arbitrarily selected entries in U', G' and C, respectively. We can add this sentinel as padding entries in the first $k-1$ k-grams of the trace (as suggested in [67]), or the last $k-1$ k-grams as discussed in Chapter 3.

As can be easily observed, the new automaton is richer in comparison to that of the Self-based IDS. In the basic Self-based IDS, the alphabet (Σ) is the the set of possible system call numbers S. In PAC-based IDS, the alphabet is now defined as a tuple $U' \times G' \times S^+ \times C$. Note that while we have focused on Unix, the categorization approach can be extended to other OSes.

4.2.2 A Simple Category Specification Scheme

We now give a simple scheme for defining the abstraction and the categories. The category specification is constructed by taking into account the importance or the sensitivity level of files/directories in the underlying OS from the security standpoint. The main goal of the specification is to separate operations which have potential security risks from the benign ones.

A fragment of an example specification is given in Figure 4.1. It consists of four sections: *euid*, *egid*, *argument categorization* and *illegal transitions* which are to be elaborated below. For each section, the category specifications are processed in the sequential manner, from the start to the end of the definition. In this fashion, the more specific

[1] As elaborated in [167], every state $q \neq q_\perp$ is an accepting state. The special state $q_\perp$ is non-accepting and contains a self-loop $q_\perp \xrightarrow{\sigma} q_\perp$ for every $\sigma \in \Sigma$. When a state q contains no outgoing transitions on $\sigma \in \Sigma$, we add an implicit transition $q \xrightarrow{\sigma} q_\perp$. Note that this FSA definition takes a single *transition mismatch*, i.e. a non-existent transition from a state $q \neq q_\perp$, as an intrusion. In practice, mainly to keep false positive rate reasonably low, the IDS raises an alarm only if the output of its anomaly-signal measurement function (e.g. Locality Frame Count) has reached a specified threshold value.

mappings can be made first and the most general ones last. This is similar to the ordering in firewall configuration files. Note that this example is only meant to be illustrative on how the category specification works. A more complete example specification can be seen in the Appendix A.

```
# EUID Abstraction Section
# Format:  <categorized-euid>:<euid1>,<euid2>,...
0:0
1:2000,2001,2003
100:*
# EGID Abstraction Section
# Format:  <categorized-egid>:<egid1>,<egid2>,...
0:0
1:1,2,3,5
100:*
# Argument Abstraction Section
# Format:  <syscall> <arg1> <arg2> <arg3> ...  <cat-value>
open p=/etc/passwd o=O_WRONLY|o=O_RDRW * 1
open p=/etc/shadow o=O_WRONLY|o=O_RDRW * 2
open * * * 18
chmod p=/etc/{passwd,shadow,group,hosts.equiv}|p=/proc/kmem * 1
# Illegal Transitions Section
# Format:  <syscall> <cat-values> [<cat-euid>,<cat-egid>]*
open [1..6,8-11,13-15] 0,* *,0
{chmod,fchmod,chown,fchown,lchown,mknod,unlink,init_module,execve} 1 0,* *,0
```

Figure 4.1: An example of category specification for the PAC-based IDS.

Privilege Abstraction Section

The euid and egid sections are meant to provide the actual value mapping for $EuidCat : U \rightarrow U'$ and $EgidCat : G \rightarrow G'$. The example specification uses the following syntax for euid and egid:

$$\begin{aligned} \langle u'_i \rangle &: \langle u_{i1} \rangle, \langle u_{i2} \rangle, \ldots, \langle u_{in} \rangle \\ \langle g'_i \rangle &: \langle g_{i1} \rangle, \langle g_{i2} \rangle, \ldots, \langle g_{in} \rangle \end{aligned} \tag{4.1}$$

where $u'_i \in U'$, $u_{ij} \in U$, $g'_i \in G'$ and $g_{ij} \in G$.

To ensure that $EuidCat/EgidCat$ is a total mapping, a special entry "*" is employed to indicate other (remaining) euids/egids so that the mapping satisfies the requirement for a function. As euid=0 and egid=0 signify important privileges in Unix, each of them has a distinguished mapping.

Argument Abstraction Section

The specification is a straightforward one. It maps the system call together with its corresponding arguments into a number representing its category. We now briefly discuss some considerations in creating a definition:

- The approach we use is to focus the specification on a subset of system calls $S' \subset S$ which should be checked in order to prevent attacks aimed at gaining full control of the system. Our choice for the system call subset S' is based on the work of Bernaschi et al. [21] which classified Unix system calls according to their *threat level* with respect to the system penetration. Here, we consider S' to be the system calls in *Threat-level 1 Category*, namely: `open`, `chmod`, `fchmod`, `chown`, `fchown`, `lchown`, `rename`, `link`, `unlink`, `symlink`, `mount`, `mknod`, `init_module` and `execve` [21].

  Other system calls in $S - S'$, which have not been defined in the specification are mapped to a unique default value. We do not address the issues raised by the system calls in Threat-level 2 (which can be used for a denial of service) and Threat-level 3 (which can be used for subverting the invoking process). Otherwise, we might need a richer IDS model which also deals with issues such as memory/storage consumption metering and file usage pattern, which are beyond the scope of this work. One advantage of the system call subset, which is approximately 10% of the total number of system calls, is that it reduces the monitoring overheads which is important when the IDS is run on-line.

  Bernaschi et al. also grouped the setuid/setgid system call family into the Threat-level 1 list. However, we take a different approach here by capturing the effect of the setuid/setgid system call family as *changes in process credential values* —in the form of (euid, egid) pairs— to form part of the state information in our enhanced IDS model.

- Given a system call $s' \in S'$, a simple approach for the choice of abstraction is to ensure that any critical operations on security-sensitive objects are mapped to a value different from non or less critical ones.

- Pathnames require special treatment and we use a special notation: `p=`<*pathname*>. Because pathnames in Unix are not unique, they have to be made canonical by turning them into a normalized absolute pathname (see also [125]).

- Additionally, we also make use of a special tag `o=`<*option-token*> to indicate that an *option-token* is to appear as an entry in a list of entries in the corresponding argument space. An example is `oflag` as the second argument of `open()`, where `O_RDWR` can co-exist with other modes such as `O_CREAT` or `O_APPEND`. A categorization specification entry is triggered if the *option-token* is present in the list.

4.2.3 Disallowing Transitions

It is also useful to specify the transitions that can immediately lead to "bad states". The idea is to identify those singleton system calls with the corresponding privileges which can

be sufficient to compromise the system's security. An example would be the operation of `chown()` on `/etc/passwd` with root privileges. Thus, the usual way of measuring anomaly signal by means of Locality Frame Count (LFC) function as used in [140] is inadequate. This can also be used as an enhancement to the access control to actually deny such a system call invocation in a running program.

In category specification, e.g. Figure 4.1, the syntax for bad transitions is:

$$s' \quad c \quad [u', g']^* \tag{4.2}$$

where c is the abstracted value for the arguments of system call s'; whereas u' and g' are the abstracted user and group privilege respectively. We denote the set of bad transitions as D_0. Disallowing D_0, however, may be too strict, and needs to be adjusted with respect to the normal traces. When extracting the normal profile from the normal trace, we can construct the set D_N from all $\sigma \in D_0$ which are also present in the normal traces. The final adjusted negative transitions are: $D = D_0 - D_N$.

In the detection stage, the PAC-based IDS flags any system call execution which matches an illegal transition $\sigma_d \in D$ as intrusive.[2] Since the PAC-based IDS can function as an IPS, it may also prevent such operation from being executed. Hence, the anomaly-signal measurement function in use must treat these bad transitions differently in determining the anomaly signal.

4.3 Experiments on PAC-based IDS

Our aim here is to evaluate whether the inclusion of abstracted process privilege and system call arguments in the PAC-based IDS can make it more resistant against mimicry attacks. In addition, we also would like to evaluate the behavior of the PAC-based IDS, particularly whether there is any increase in its false positives.

4.3.1 Attack Construction on PAC-based IDS

We extend the attack construction algorithm on the Self-based IDS (described in Section 3.2.5) to also work with the PAC-based IDS. This can be easily done since we just need to extend the k system call labels in each node's state in the overlapping graph G by adding the euid, the egid and the argument category value.

To compare the PAC-based IDS with the two variants of the Self-based IDS analyzed earlier, denoted as SET-IDS and GRA-IDS (see Section 3.3), we evaluate the attack construction on the PAC-based IDS using the same three vulnerable programs used in evaluating SET-IDS and GRA-IDS. These vulnerable programs are:

- Traceroute: with the basic attack trace: `open()`, `write()`, `close()`, `_exit()`.
- JOE: with the attack trace: `open()`, `write()`, `close()`, `_exit()`.

[2]The PAC-based IDS can additionally log any execution of $\sigma \in D_N$.

- WU-FTPD: with the same attack trace as in [167], which is: `setreuid()`,`chroot()`, `chdir()`, `chroot()`, `open()`, `write()`, `close()`, `_exit()`.

In the trace generation, we purposely set the euid/egid value of all system call entries in the normal trace to 0 (root). In other words, we assume a poorly written setuid program. This is to provide the worst-case condition for attacks to occur, since system calls in the attack trace are typically to be executed with root privilege.

In these experiments, the argument category specification makes use of the "dangerous" system call subset discussed in Section 4.2. For the choice of arguments, from [21, Table 4], we can see that the dangerous arguments for system calls in S' are mainly files/directories. The work by Garfinkel and Spafford [51, Appendix B] gave a comprehensive list of security sensitive and important files/directories that one might want to consider monitoring in Unix. In our experiments, we pay special attention to several security critical files in the Unix/Linux environment, which are listed in Table 4.1. For simplicity, we omit most of the system configuration files in `/etc` (such as: `/etc/inetd.conf`, `/etc/hosts`, `/etc/cron/*`) and devices files in `/dev`. We however include entries for various directories commonly found in the Unix/Linux file system hierarchy, which conforms to the *Filesystem Hierarchy Standard* (`http://www.pathname.com/fhs/pub/fhs-2.3.html`). While one can use a more detailed specification, ours is already sufficient to show an increase in the robustness of the IDS.

File ID	**File name**
F_1	`/etc/passwd`
F_2	`/etc/shadow`
F_3	`/etc/group`
F_4	`/proc/kmem`
F_5	`hosts.equiv`

Table 4.1: Several important files to be protected from security viewpoint.

Based on our experiments using window sizes from k=5 to 11, we found that no stealthy attack could be constructed, both under the code-injection and the trojan attack scenarios, on the PAC-based IDS using the three sample programs which previously allowed for stealthy attack on both SET-IDS and GRA-IDS.[3]

We now experiment with the PAC-based IDS against various mimicry attack strategies, and also investigate its false positives.

[3]If no attack can be constructed on windows size i, then the algorithm cannot construct the attack on window sizes larger than i.

4.3.2 Behavior of the PAC-based IDS

Resistance Against Various Attacks

Having shown that the PAC-based IDS can better withstand mimicry attacks on the three sample programs, we now evaluate the IDS against a number of different attack strategies.

In Table 4.2, we list a number of common attack strategies in the Unix/Linux environment on the files listed in Table 4.1 when the system calls are executed with a root euid/egid privilege. While the list is not comprehensive, it is sufficient to demonstrate the improvements in the IDS' resistance level. We choose the traceroute program for this experiment. The experiment was done on normal traces described earlier (2,789 system calls) with the sliding-window size (k) set to 5, which is a relatively short one. We found that all the attack strategies listed in Table 4.2 fail on the tested normal traces even in the trojan attack scenario. For most of the strategies (A_6–A_{61}), the attacks fail because the needed attack system calls do not appear in the normal traces. Thus, these attack strategies would not work either to attack the SET-IDS and the GRA-IDS. In attacks A_1–A_5, given the category specification, the attack searches fail because the normal trace does not contain the required system call argument categories.

Attack ID	**Operation (respectively)**
$A_1 - A_5$	Open and write an entry into F_1, F_2, F_3, F_4, F_5
$A_6 - A_{10}$	Chmod on F_1, F_2, F_3, F_4, F_5
$A_{11} - A_{15}$	Fchmod on F_1, F_2, F_3, F_4, F_5
$A_{16} - A_{20}$	Chown on F_1, F_2, F_3, F_4, F_5
$A_{21} - A_{25}$	Fchown on F_1, F_2, F_3, F_4, F_5
$A_{26} - A_{30}$	Lchown on F_1, F_2, F_3, F_4, F_5
$A_{31} - A_{35}$	Rename F_1, F_2, F_3, F_4, F_5 into some other file
$A_{36} - A_{40}$	Rename some other file into F_1, F_2, F_3, F_4, F_5
$A_{41} - A_{45}$	Link F_1, F_2, F_3, F_4, F_5 into some other file
$A_{46} - A_{50}$	Link some other file into F_1, F_2, F_3, F_4, F_5
$A_{51} - A_{55}$	Unlink F_1, F_2, F_3, F_4, F_5
$A_{56} - A_{60}$	Mknod F_1, F_2, F_3, F_4, F_5
A_{61}	Execve a shell or command

Table 4.2: Attack strategies (on files listed in Table 4.1) to be prevented.

False Positives

Here, we give some preliminary results of comparing the PAC-based IDS in terms of its false positives to the baseline Self-based IDS. We choose two programs, ls and traceroute in Red Hat Linux 7.3. For each program, we record 10 traces from 10 different program runs. We then randomly choose one to be tested against the other 9. Now, we compare

the PAC-based IDS with the Self-based IDS where the normal profile is a set of k-grams (SET-IDS). The comparison results for k=5 to 11 are shown in Table 4.3. Here, we simply measure the number of foreign k-grams. As can be seen, the enhancement does not increase the number of false positives.

	traceroute		**ls**	
k	**SET-IDS**	**PAC-IDS**	**SET-IDS**	**PAC-IDS**
5	0	0	2	2
6	0	0	2	2
7	0	0	2	2
8	0	0	2	2
9	1	1	2	2
10	2	2	2	2
11	3	3	2	2

Table 4.3: Number of foreign k-grams in traceroute and ls with window sizes k=5 to 11. SET-IDS refers to the Self-based IDS (Stide) whereas PAC-IDS indicates our new PAC-based IDS model.

4.4 Discussions

We have shown that the PAC-based IDS model is more resistant to mimicry attacks since the basic attacks in our experiments could not be turned into the mimicry attacks.

Furthermore, we have the following observations:

- As we have shown that the Self-based IDS with certain sizes of the sliding window, such as k=6 to 8 as suggested in [60, 140, 67], is insufficient, other improvements to the IDS model are thus necessary. Our privilege and argument abstraction techniques seems to be able to answer the need to make the Self-based IDS more robust. In addition, one can always specify his/her own specification rules for the PAC-based IDS in order to suit a particular program in preventing possible attacks.
- Our experimental results show that with the given basic attacks, it was not possible to turn them into the mimicry attacks on the PAC-based IDS although it was possible to do so in the two Self-based IDS models (SET-IDS and GRA-IDS). Most results that we are aware of for analyzing an IDS, in particular using the mimicry attacks, are usually of the negative variety in that they show the existence of attacks or ways of attacking the IDS. It is significant that our result here is a positive one, since it shows that certain systematic attacks fail to work on the PAC-based IDS. However, we do not guarantee that no attacks are possible since the evaluation is relative with respect to a given basic attack and the normal traces. The question of a security guarantee is in fact an open problem in the most IDS models, and we argue that our work here paves a way towards the more robust evaluation methods.

- The false positive results of the experiments are encouraging as we find that improving the IDS with a more fine-grained detection mechanism does not increase the number of false positives over the baseline IDS.

- The privilege and argument abstraction technique can also be applied to the other gray-box IDS models, such as the FSA-based model [136]. In this new model, we have the set of states $\mathcal{Q} = \{q_0, q_\perp\} \cup \{U' \times G' \times P\}$ with $P =$ set of possible Program Counter values, and $\Sigma = S \times C$. The transition is thus enhanced using a 2-tuple of the system call number and the argument category value.

4.5 Summary

We have proposed an extension to the Self-based IDS using privilege and argument abstraction. The new IDS, called the PAC-based IDS, is a more-fine grained model which takes into account the security aspects of the operations. We argue that this extension is both simple to use and also makes the IDS more robust. Our experimental results show that mimicry attacks, which were able to work in the baseline Self-based IDS, fail in the PAC-based IDS. We also have some evidences that the increase is detection accuracy does not lead to more mis-predictions.

An important advantage of our IDS extension is its *simplicity*. Directly using the arguments or process credentials will not work well due to the possibility of increasing the number of false positives. However, a simple classification which abstracts away the irrelevant information and takes into account a security model of the OS does work. Furthermore, the simplicity means that the scheme can be easily integrated into various IDS models to make them significantly more secure.

Chapter 5

Lightweight Executable Integrity Protection

Ensuring the integrity of an executable program is an important and essential step in the Program Protection Life Cycle (PPLC). It helps establish a stronger trust on good program execution. It is also beneficial to the OS since it can help protect the OS against malicious kernel drivers, i.e. attacks which attempt to load malware into the kernel. Section 2.3 has provided some background on the executable integrity problem, as well as briefly mentioned existing authentication systems and issues in providing the executable integrity on Windows. For ease of reference in this chapter, we refer to an executable code stored in the file system as a *binary*.

Despite its important role, binary integrity protection mechanism is yet to be commonly incorporated as a standard in-kernel mechanism in popular commercial-off-the-shelf OSes, such as Windows and Solaris. In this chapter, through a proof-of-concept implementation called BinAuth, we demonstrate the practicality of establishing a binary authentication system on Windows. BinAuth is a practical, lightweight and in-kernel mandatory binary authentication system, designed to allow the execution of only trusted binaries on a host. It is lightweight in terms of both performance overheads and its reduced reliance on PKI for the certificate revocation service.

Demonstrating the practicality of such an enforcement system for trusted binary execution on a complex commodity OS is significant. This is since many of the security problems on commodity OSes, particularly Windows, stem from a host's inability to distinguish between trusted and distrusted software. BinAuth provides a mandatory authentication for the full range of binaries under Windows, and goes beyond code authentication in Windows XP and Vista. It also protects driver loading, thus offering an increased kernel protection. Our benchmarking shows that the overhead of comprehensive binary authentication can be quite low, around 2%, with a caching strategy.

We also develop a framework for comparing binary authentication systems based on

a number of key design factors. We use this framework to compare a variety of existing authentication systems, and identify the desirable properties of an robust and efficient binary authentication system. In fact, we specifically design BinAuth to meet these desirable properties. Besides this framework, we additionally propose a simple scheme called Software_ID which leverages on the authentication infrastructure. This scheme is similar to the Common Platform Enumeration (CPE) initiative [104], which can help simplify the task of vulnerability alert management as discussed more in Chapter 6.

Most of the results reported in this chapter have been published in our collaborative works [59, 173]. The author of this thesis focuses mainly on the design of BinAuth including the notion of Software_ID, and on a comparative analysis of the authentication schemes. The implementation of BinAuth on Windows was done by other team members. The remainder of this chapter is organized as follows. We first give some definitions of authentication goals in Section 5.1. Section 5.2 outlines our framework for comparing binary authentication systems. Section 5.3 then describes the design, and briefly mentions some implementation aspects of BinAuth. We analyze the security of BinAuth in Section 5.4. We then provide the benchmarking results on BinAuth in Section 5.5. Section 5.6 discusses the Software_ID scheme which can take advantage of BinAuth. Finally, Section 5.7 summarizes this chapter.

5.1 Definition of Authentication Goals

Let us first establish some definitions on binary authentication goals. Note that throughout this chapter, we define "*administrator*" to be the trusted privileged user, such as system administrator in Windows or superuser in Unix. The goal of a binary integrity protection system is to ensure that an executed binary only comes from a trusted software developer, and that it is executed in the correct execution context. To be more specific, we define the security goals of binary authentication (on stored executables) as follows: "A binary authentication system aims to ensure that a stored binary file B is allowed for execution at time $t_{invoked}$ only if the following *authentication goals* are satisfied":

- (G_1) ***Trusted-Origin***: it comes from a trusted source, which was verifiable or was deemed trusted by the administrator at time $t_{installed}$, i.e. when the binary was added (installed) into the system; and
- (G_2) ***Unmodified-Binary***: it remains unchanged from the time $t_{installed}$.

Since binaries are kept in files and that they are identified by their pathnames on the host's file system, goal G_2 can be further refined into the following *integrity guarantees*:

- (A_1) ***Binary-Content Integrity***: the *content* of a binary file is unmodified since time $t_{installed}$. This ensures that a binary has not been tampered with. For example, `cmd.exe` in a Windows system is not replaced with a trojan.

- (A_2) ***Binary-Pathname Integrity***: the pathname (location) of a binary file must match its content, based on the pathname mapping when it was securely added into the host at $t_{installed}$. This ensures that we are executing an executable which we actually want. For instance, suppose that the contents of a file-system format and a shell binaries are both authenticated. However, if an attacker swaps their pathnames, then running the shell would cause the file system to be formatted.

5.2 Framework for Analyzing Binary Authentication Schemes

This section describes our framework for characterizing authentication systems based on a number of key design factors. We elaborate the available choices for the design factors, and discuss their impact on the effectiveness of the resultant systems. A set of desirable design factors are then identified. We also compare a variety of existing authentication systems using the framework.

5.2.1 Security Assumptions

First, we make clear of some assumptions which are commonly made by software-based binary authentication systems including BinAuth:

- **Trusted host's kernel**: Since the authentication systems run on top of the kernel, and thus depend on the kernel to achieve their security objectives, it is assumed that the host's kernel is uncompromised. Following a host (but not kernel) breach, an authentication system is still expected to remain trusted and to function properly in preventing the execution of illegal binaries. As such, an authentication system should allow for the possibility of malware exploitation, e.g. a buffer overflow attack hijacking a privileged process, but not the ability to alter the kernel code [174]. There exist authentication systems which still can function despite a kernel breach, such as [122]. A system like this, however, usually requires additional hardware and software which are deployed independently from a host.
- **Trusted authentication information**: The systems also assume the security (integrity, confidentiality and availability) of all the keys used for authentication as well as any database and configuration files stored on the host's file system.
- **Trusted administrator account**: Since the administrator in most OSes virtually has unblocked access to all the files, registries, and possibly kernel settings, we also need to assume the security of the administrator account.

5.2.2 Authentication System Design Options

We now list the main design options in constructing a binary authentication system.

(D_1) **Location of the authentication/verifier module**:

- **(a)** *In-kernel module*: The integrity-verifier module is implemented as part of the kernel, and thus runs in the kernel space.
- **(b)** *User-mode application*: The module runs as an application in the user space.

(D_2) **Time of authentication**:

- **(a)** *Pre-execution*: The binary file is verified *just right before* it is loaded for execution. This does not necessarily imply that the authentication is done by in-kernel module since wrappers can be placed on various entry points of the program's execution, e.g. command shell and GUI-based shell.
- **(b)** *User-specified time*: The authentication timing is set independent from the binary-invocation time. It can be directly triggered by the user, or based on a user-specified time in the case of a scheduled job. A system like this usually verifies all protected binaries within one verification session. It thus takes a "snapshot" of all the binaries at one point in time.
- **(c)** *Installation-time*: The verification is performed only at the installation time.

(D_3) **Coverage of mandatory authentication**:

- **(a)** *All binaries*: All binaries must be verified before being executed.
- **(b)** *Selected binaries*: Only certain classes of binaries are subject to verification. For example, only binaries run with the administrator privilege and those with embedded digital signatures are checked. The rest are allowed to run without authentication.
- **(c)** *Configurable*: Here, a system's level of enforcement is configurable. The administrator is responsible to select one from a set of preset options to apply at a time. He/she, for instance, can apply one of the following options: all binaries must be authenticated, all binaries with the administrator privilege are authenticated, or no authentication is performed.

We note that for the mandatory integrity enforcement on all binaries, an in-kernel pre-execution mechanism is the most robust option.

(D_4) **Implementation platform**:

- **(a)** *Microsoft Windows*
- **(b)** *Unix/Linux*

We note that it is theoretically possible to port an authentication system implemented in one OS to another OS. In practice, however, such porting is rather difficult due to the differences in the security model and the kernel operations of the underlying OSes.

(D_5) **Authentication services provided**:

(a) *Binary-Content Authentication*: as defined as A_1 (in Section 5.1).

(b) *Binary-Pathname Authentication*: as defined as A_2 (in Section 5.1).

(D_6) **Placement of the integrity information (signature/hash)**:

(a) *Embedded into the executable*: In this method, the signature or hash value is embedded into the executable file. Putting the information in the file seems to result in a cleaner model of the authentication system. However, as we will see later there is a potential weakness in such a system.

(b) *(Secure) centralized database*: The information from all the protected binaries is stored in a database. If all binaries are to be mandatorily authenticated prior to execution, then the database serves as a "whitelist" of accepted applications. This contrasts with the "blacklist" approach taken, for example, by anti-virus software. The advantage of using a centralized database is that the authentication system can also monitor non-binary files.

(D_7) **Producer of the binaries' signatures/hashes**:

(a) *Authentication system*: Here, all the hashes/signatures come solely from the authentication system. The administrator assumes that a binary file is good at some point in time (usually at $t_{installed}$), and then invokes the authentication system to produce a signature/hash as the reference for future executions.

(b) *Developer*: All binaries are assumed to be signed by their respective software developers.

(c) *Both*: It is possible that the administrator produces the signatures/hashes on some of the binaries. It is also possible that the administrator reproduces the signatures/hashes based on the existing developer's signature for security or performance reasons as proposed in our BinAuth system.

(D_8) **Authentication mechanism used**:

(a) *(Keyed) hash function*: The integrity of a binary is checked by using a (keyed) hash function.

(b) *Digital signature operation (with PKI)*: The checking is done by using digital signature operation and related PKI mechanisms.

(D_9) **Authentication Caching**:

(a) *With/without caching*: Whether a caching technique is employed by an in-kernel authentication system to keep track of the previously authenticated binaries. This technique is employed to reduce the overheads of always checking the binary files every time they are executed. It is particularly useful for some frequently executed binaries, especially if their file sizes are large. Files which are stored externally, e.g. files on NFS, however cannot be cached [8]. One

main drawback of the caching technique is that the authentication system now needs to monitor if a binary in the cache list has been modified.

(b) *Not-applicable*: The caching technique is not applicable to a user-mode authentication system, or one such as [174] which checks the integrity during binary installation.

(D_{10}) **Reduced reliance on PKI**:

(a) *Reduced/full reliance*: Whether there is any mechanism employed to reduce the reliance on PKI in the case where digital signatures are used.

(b) *Not-applicable*: This is not applicable if only hash function is used.

(D_{11}) **Support for secure update of a protected binary**:

(a) *With/without support*: Whether there is any mechanism to support secure update of a protected binary file.

5.2.3 Comparison of Existing Methods and BinAuth using the Framework

We now survey and compare various existing authentication systems. We select a representative system for each category. We also highlight how well each authentication scheme meets the authentication goals G_1 and G_2, and contrast them with BinAuth. Table 5.1 summarizes the surveyed systems including BinAuth.

Tripwire [74] is one of the first schemes to perform file integrity protection. Tripwire is however limited as it is a user-mode application program, and checks the file integrity in an off-line manner. It also does not provide any mandatory form of integrity checking. In addition, there exist many known attacks such as: file modification in between authentication times, and attacks on system daemons (e.g. `cron` and `sendmail`) and system files that it depends on [11].

There are a number of in-kernel binary authentication implementations such as DigSig [8], Trojanproof [170], SignedExec [162] and one proposed in [28]. These are mainly for Unix or Linux. They modify the Unix kernel to verify an executable's digital signature prior to its execution. **DigSig**, **SignedExec** and the scheme proposed in [28] embed signatures within the binary itself by making use of the ELF format. It appears that these system provide binary-content authentication (A_1), but not binary-pathname authentication (A_2). The problem is that, in an embedded-signature authentication system, one signed binary replacing another signed binary would go undetected. One solution to this problem is that the authentication system must include the pathname in addition to the binary content when producing a digital signature. Even with this inclusion, the system would still suffer from a "old-attack" problem, when the attacker replaces a signed binary with an older, perhaps vulnerable, signed binary of the same pathname. DigSig

Authentication System	Verifier Module Location	Time of Authentication	Enforced Mandatory Authentication?	Platform	Authentication Provided		Placement of Integrity Information	Signature /hash Producer	Mechanisms Used	Caching Used?	PKI Reliance Reduction?	Binary Update Support?
					Binary Content	Binary Pathname						
Tripwire	Application-level	User-specified	No	Unix/Linux, Windows	Yes	Yes	Centralized	Auth. System	Hash	No	-N/A-	No
DigSig	In-kernel	Pre-execution	No	Linux	Yes	No	Embedded	Auth. System	Hash	Yes	No	No
Trojan-proof	In-kernel	Pre-execution	Yes	Linux	Yes	Yes	Centralized	Auth. System	Hash	No	-N/A-	No
SignedExec	In-kernel	Pre-execution	Yes	Linux	Yes	No	Embedded	Auth. System	Digital Signature	Yes	No	No
CryptoMark	In-kernel	Pre-execution	Configurable	Linux	Yes	No	Embedded	Auth. System	Digital Signature	No	No	No
Emu System	In-kernel	Pre-execution	Yes	Windows	Yes	Yes	Centralized	Auth. System	Hash	No	-N/A-	No
SignTool	Application-level	User-specified	No	Windows	Yes	No	Embedded	Auth. System	Digital Signature	No	No	No
Sigcheck	Application-level	User-specified	No	Windows	Yes	No	Embedded	Auth. System	Digital Signature	No	No	No
Vista UAC (Signed Executables)	In-kernel	Privilege-elevation	No	Windows	Yes	No	Embedded	Developer	Digital Signature	No	No	No
Wurster-Oorschot	In-kernel	Installation time	No	Unix/Linux	Yes	No	Embedded	Developer	Digital Signature	-N/A-	Yes	Yes (weak)
BinAuth	In-kernel	Pre-execution	Yes	Windows	Yes	Yes	Centralized	Auth. System, Developer	Hash, Digital Signature	Yes	Yes	Yes

Table 5.1: Comparison of binary authentication systems using the design-option based framework.

employs a technique called signature revocation list in order to blacklist a signed binary. Catuogno and Visconti similarly proposed a file revocation list and an alternative tree-based scheme [28]. However, such techniques require a separate centralized configuration file which can grow over time. Maintaining such a list also poses its own challenges. In light of this, we opt for an authentication system with a centralized database.

For efficiency, DigSig and the scheme proposed in [28] employ a caching mechanism to avoid checking binaries which have been previously verified. In developing BinAuth, we also address the detailed implementation issues of caching technique on Windows, which are much more complex than in Unix.

Cryptomark [18] is another in-kernel authentication system on Linux, and is similar to DigSig in a number of ways. The special feature of Cryptomark is that it can be configured to require valid digital signatures for *all* or *some* binary files. The most secure configuration requires a valid signature for every binary. A more permissive configuration mandates signatures only for binaries that run with the root privilege.

There are some mechanisms in Windows related to binary authentication. **Authenticode** [55] is a Microsoft infrastructure for digitally signing binaries. In Windows versions prior to Vista, such XP with SP2, Authenticode is used as follows:

- During ActiveX installation: Internet Explorer uses Authenticode to examine the ActiveX plugin, and shows a prompt which contains the publisher's information including the result of the signature verification.
- When a user downloads a file using Internet Explorer: If this file is executed using the Windows Explorer shell, a prompt is displayed giving the information of the publisher's information. Internet Explorer uses an NTFS feature called Alternate Data Streams to embed the untrusted Internet zone information into the file. The Windows Explorer shell detects the zone information and then displays the prompt. This mechanism is *not* mandatory and relies on the use of zone-aware programs.

Since Authenticode runs in user space, it can be bypassed in a number of ways, e.g. from the command shell. It is also limited only to files downloaded using Internet Explorer. Only the EXE binary is examined by Authenticode, but the DLLs are ignored. One possible attack is to put malware into a DLL and then execute it, e.g. with `rundll32.exe`. Furthermore, Authenticode relies heavily on digital certificates. Checking CRL may add extra delay including timeouts due to the need to contact the CA and download the latest CRL data. In some cases, this causes significant slowdown.

Windows Vista improves on signed checking since its **User Account Control (UAC)** can be configured for mandatory checking of signed executables. However, this is quite limited since the UAC mechanism only kicks in when a process requests privileged elevation, or for certain operations on protected resources. Vista does not seem to prevent the loading of unsigned DLLs and other non-EXE binaries. The 32-bit versions of Windows

(including Vista 32-bit) do not check whether drivers are signed. Similar to DigSig, the UAC as well as Authenticode do not authenticate the binary pathname. Moreover, a system that always requires PKI infrastructure, we believe, poses various challenges for a general purpose online (pre-execution) integrity checking mechanism.

One of the existing works that is closest to our BinAuth scheme is the **Emu** system [132] which also runs in Windows. It intercepts a process creation by intercepting the `NtCreateProcess()` system call. It is unclear, however, whether they are able authenticate all binary codes since trapping at `NtCreateProcess()` is insufficient to deal with DLLs. No performance benchmarks were given, so it is difficult to assess its efficiency.

There is also a binary protection system which protects unauthorized modification of existing binaries during the installation of a new piece of software [174]. This system does not restrict the software which can be installed on a system, but it denies one which modifies existing binaries. Hence, the system does nothing to prevent malware from being installed on the system, but it restricts the files that the malware can modify. One particular feature of the system in [174] is that it aims to allow software updates in a controlled manner. A binary can be replaced by another binary only if the latter is of the same name and contains a digital signature derived from the same public key previously used on the former. However, there is a potential attack of downgrading the binary. BinAuth, in contrast, makes use of a naming scheme like Software_ID (described in Section 5.6) to ensure a controlled software upgrading.

Lastly, we remark that there was also a related work which compares existing authentication systems. In [110], Motara and Irwin looked at six authentication systems and evaluated them in terms of the following criteria: ease of use, executable checking, pre-execution validation, active development and transparent checking. They also discuss several design options, namely: in-kernel vs user-mode (D_1), pre-execution vs offline validation (D_2), embedded-signature vs centralized (D_6), and caching mechanism (D_9). Our work [173] conducts a significantly more comprehensive analysis on the design options when compared to [110]. More importantly, our analysis is performed as a part of our effort to devise a robust yet lightweight binary authentication system. As such, the analysis is geared towards identifying desirable characteristics for such an authentication system, which are then realized in the BinAuth system.

5.3 System Architecture for Lightweight Authentication

Based on the analysis using design-option based framework discussed above, our goal is to devise BinAuth as a robust and practical authentication scheme with the following characteristics: an in-kernel (D_{1a}), pre-execution (D_{2a}), mandatory (D_{3a}) authentication scheme for Windows (D_{4a}), ensuring both binary-content (D_{5a}) and binary-pathname

integrity (D_{5b}) by storing the integrity information in a centralized database (D_{6b}). BinAuth accepts both digital signatures from the developer as well as produces MAC values on protected binaries. It thus exercises (D_{7c}), (D_{8a}) and (D_{8b}), and has a reduced reliance on PKI ($D_{10a-Reduced\ reliance}$). It also makes use of caching ($D_{9a-With\ caching}$), and supports controlled update of the protected binaries ($D_{11a-With\ support}$). Additionally, it should incur low overhead and allow for a better management of binaries.

5.3.1 BinAuth Architecture

In the following discussion, we assume that binaries already come tagged with a necessary naming information such as Software_ID (see Section 5.6) or CPE [104][1]. During the binary authentication setup, preferably done immediately after the targeted binary installation, we generate the Message Authentication Code (MAC) value for each binary. In the case where a binary is digitally signed by its developer, we verify the signature first before generating the MAC value. Hence, *only one* public-key operation is done per signed binary. We choose to use a keyed hash construction, the HMAC algorithm [78], so that there is a secret key for the administrator. This secret key could be stored on a secure external storage, e.g. a TPM module [160]. Given this secret key, the execution of an illegally-added binary by an attacker who is also able to modify the centralized database will still be detected. To authenticate binary integrity for any future execution of the code, only the generated HMAC value needs to be checked. In what follows, we mostly write MAC which already covers the choice of HMAC.

One way of storing the generated MAC is to embed it into the binary. However, doing so may interfere with file format of the binaries, and may also have other complications, such as the inability to provide a binary-pathname authentication guarantee (A_2). As such, we instead use a repository (database) file which stores all the MAC values of authenticated binaries together with their pathnames.[2] During the boot-up process, the kernel creates its own in-memory data structures for binary authentication from this repository. We can also customize binary authentication on a per-user basis rather than on a system-wide basis, thus producing a white list of binaries approved for execution for each user. In the case when the initial binary does not have a digital signature, then the administrator can choose to approve the binary and generate a MAC for it.

There are two main components of the system: the **SignatureToMac** and the **Verifier**. The SignatureToMac maintains the authentication repository, called `Digest_file`. The Verifier is a kernel driver which makes use of `Digest_file` and decides whether the execution of a binary is to be allowed.

[1] Unlike our Software_ID, CPE does not include the binary file name. Hence, the file name must be added should the CPE be used.

[2] We assume that the repository file is protected from illegal modifications.

5.3.2 SignatureToMac

Once a piece of software is installed on the system, Figure 5.1 shows how SignatureToMac processes the installed binaries. The steps involved can be elaborated as follows.

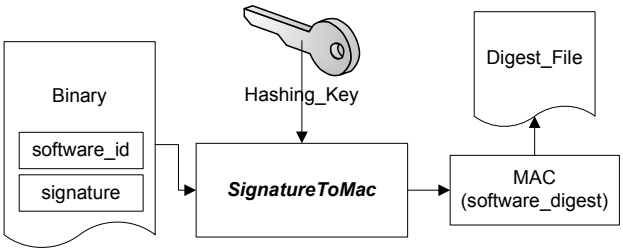


Figure 5.1: SignatureToMac: deriving the MAC for a signed binary.

1. If a binary is signed, check the validity of the binary's digital signature and corresponding certificate. If the signature or certificate is invalid, report failure.
2. Consult the administrator whether the software is to be trusted or not. This is similar to the Vista UAC dialog, but only happens once. Additional checking policies (possibly mandatory) can also be implemented.
3. Generates the MAC value (called `software_digest`) from the binary's content (including its `Software_ID` string) using a system-wide secret key called `Hashing_key`. This `Hashing_key` is accessible only by the authentication system, e.g. obtained at boot time from a secure (external) storage.
4. Add an entry for the binary as a tuple ⟨`path_name`, `software_digest`⟩ into the `Digest_file` repository, and inform the Verifier to perform the necessary in-memory updates.

5.3.3 Verifier

The Verifier performs a mandatory binary authentication — it denies the execution of any kind of Windows binary which fails to match the MAC and the pathname. There are two general approaches for the checking. One is *cached MAC*, which avoids generating and checking the MAC for a previously authenticated binary. The other is *uncached MAC*, which always generates and checks the MAC. As we will see, these two approaches have various tradeoffs. The cached MAC implementation needs to ensure that binaries are unmodified. Hence, the Verifier monitors the usage of previously authenticated binaries on the cache, and removes them from the cache if they can potentially be modified.

The core data structure of the Verifier component can be viewed as a table of tuples in the form ⟨`Kernel_path`, `FileID`, `MAC`, `Authenticated_bit`⟩ representing the allowed binaries. The table is indexed on `Kernel_path`+`FileID` to allow for a fast lookup. The four fields in the tuple are described below.

- The `Kernel_path` is Windows kernel (internal) pathname representation of a file. In Window's user space, a file can have multiple absolute pathnames, due to: (i) 8.3 file naming format (e.g. "`C:\Program Files\`" and "`C:\progra~1\`" are the same); (ii) symbolic links; (iii) hard links; (iv) volume mount points; or (v) the `SUBST` and `APPEND` DOS commands. The `Kernel_path` is a unique representation for all the possible pathnames. When the system loads a tuple ⟨`path_name`, `software_digest`⟩ from `Digest_file` during the startup, `path_name` is converted into `Kernel_path` since all subsequent checks by the Verifier use the latter.
- The `FileID` is a pair of ⟨`device_name`, `NTFS_object_ID`⟩. The `device_name` is a Windows internal name to identify a disk or partition volume. For instance, the device name `HarddiskVolume1` usually refers to `C:\`. The `NTFS_object_ID` is a 128-bit number uniquely identifying a file in the file system volume. The Verifier uses the `FileID` to identify the same file given more than one hard link. This prevents an attacker from creating a hard link in order to modify a binary without invalidating the binary cache. If the FAT file system is used instead of NTFS, we employ the pathname to identify a binary.
- The `MAC` is same as the `software_digest` entry in `Digest_file`. Our prototype implements the HMAC-MD5 [78], HMAC-SHA-1 and HMAC-SHA-256 [41] hash algorithms.[3]
- The `Authenticated_bit` remembers whether the binary has been previously authenticated. It is initially set to false, and then set to true after a successful authentication.

Figure 5.2 shows the authentication steps performed by the Verifier (with caching technique used) when a binary is invoked for execution. The steps can be elaborated as follows.

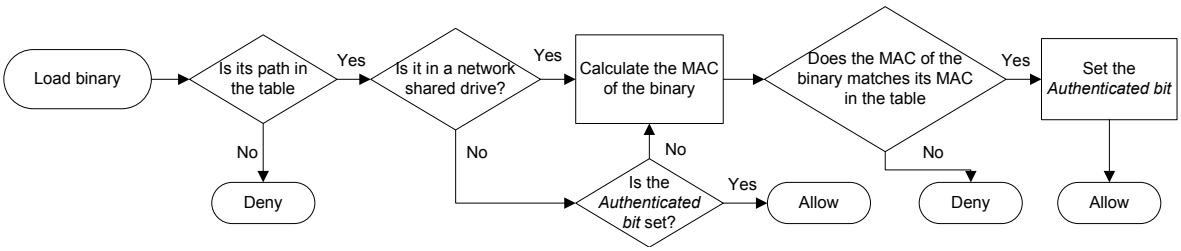


Figure 5.2: Verifier: the in-kernel authentication process.

1. Check if the binary's `Kernel_pathname` exists in the Verifier's table. If not, deny the execution (and optionally log the event). A notification is accordingly sent to the user. The user then could perform the step of generating the binary's MAC using the SignatureToMac.

[3]The stronger hash functions SHA-1 and SHA-256 are also used due to recent concerns on the weaknesses of MD5 and the associated attacks on it [168].

2. If the file is on a network shared drive or removable media, then goto step 4.
3. If the `Authenticated_bit` is set, then go to step 7.
4. Performs MAC generation operation on the binary.
5. If the resulting MAC value does not match with the `MAC` stored in the Verifier's table, deny the execution.
6. Set the `Authenticated_bit` of the binary to true.
7. Pass the control to the kernel to continue the binary's execution.

For additional details on the implementation aspects of BinAuth, such as the use of system-call interception mechanism and file-monitoring mechanism for secure caching, the readers can refer to [59, 173].

5.4 Security Analysis

In this section, we give the security analysis of BinAuth.

The security of a binary authentication system also relies on the strength of the chosen hash functions (i.e. MD5, SHA-1, SHA-256) as well as the HMAC construction. Here, we assume that any change in a binary can be detected through a changed MAC value.

When authenticating binaries with digital signatures, subsequent invocations only need an MAC verification to ensure the authenticity of the binary. In other words, MAC authentication "preserves" the previously established data integrity and data-origin authentication property of a binary from a valid digital signature. A subtlety arises when the certificate expires or is revoked at some point in time after time $t_{installed}$ (the MAC generation). We view that the question of whether one should keep trusting the binary for execution depends on one's level of trust on certificate expiration/revocation. If the certificate expiration or revocation means that the corresponding public key *must no longer be used*, but the fact that the *previously established goodness properties of a binary* still hold, then we can keep trusting the binary (as long as we still believe in its issuer).

We can reason this issue more concisely using a BAN Logic-based reasoning [25, 26]. An overview of BAN Logic is given in Section 2.6, and its extension is the subject of Chapter 8. Since we need to deal with a belief established in the past, we thus need to incorporate a temporal aspect into BAN Logic.[4]

For our purpose, we can use a temporal construct defined in [149], namely $\Diamond$, which means "*at some point in the run prior to the current one*". Given a formula φ, a run r, and a time t, we define that:

$$(r,t) \models \Diamond\, \varphi \;\; \text{iff} \;\; (r,t') \models \varphi \;\; \text{for some} \;\; t' < t. \tag{5.1}$$

[4]This is because the original BAN Logic only allows the promotion of (past) "once said" to (present) "believe" using the Nonce-verification Rule (i.e. R_2 in Appendix C).

Given this $\Diamond$ construct, we can restate the validity of the employed MAC authentication in BinAuth as follows:

> The question of whether we can still believe in a binary which is successfully verified using the MAC authentication despite the fact that the public/private key pair used to generate the digital signature has expired or has been revoked (at one point in time after the MAC generation), is equivalent to that of determining whether the following *Carried-Forward-Belief* Rule is applicable:
>
> $$\frac{P \models Q \Rightarrow X,\ \Diamond(P \models X)}{P \models X}. \tag{5.2}$$

That is, if we know that $(r, t') \models (P \models X)$ for some $t' < t$, with t being the present time, and that $(r, t) \models (P \models Q \Rightarrow X)$, we are to determine whether we can derive $(r, t) \models (P \models X)$. For our purpose of authenticating a binary, we chose to accept this rule. As such, BinAuth can verify the integrity of a binary with low performance overhead.

Note however that, given an invalid public/private key pair (expired or revoked), we cannot derive the *present* belief of a binary $(P \models X)$ using our MPKI-BAN Logic (defined in Chapter 8). This is since the New Certificate-Validation Rule (R_{15} in Appendix D) cannot be used to derive the goodness of the public $(\wp\kappa\,(P, K_P))$ and private key $\left(\Pi(K_P^{-1})\right)$. As a result, we cannot use the New Message-Meaning Rule for signed message (R_{13}) to derive $P \models Q \mid\sim X$, which is a requirement for the desired belief $P \models X$.

One additional issue with regard to the use of MAC instead of the developer's digital signature is the potential delay in obtaining a timely certificate revocation information, which may happen in practice. In SignatureToMac, BinAuth needs to ensure the validity of the developer's certificate (at time $t_{installed}$). However, it may be the case that the certificate has been already revoked by that time, but the revocation notification fails to be received in time. To overcome this, we stipulate BinAuth to monitor the status of a recently used certificate until the possibility of such a delay diminishes. That is, BinAuth will still ensure that the certificate is still valid at time $t_{ensured} > t_{installed} + \delta_{rev\_delay}$, where δ_{rev_delay} is the specified upper-bound margin for the delay in obtaining revocation information. This margin can be set in accordance with the revocation scheme's inherent timeliness delay (see Section 2.5). BinAuth can thus maintain a list of certificates to be revalidated, which are then checked periodically in the background.

We also have analyzed possible system (OS-related) attacks to BinAuth. These attacks mainly attempt to target the caching system, i.e. the attacker attempts to modify an already authenticated binary without causing the `Authenticated_bit` to be set to false. The analysis on how BinAuth can deal with these attacks can be found in [59, 173].

Lastly, we remark that if an additional hardware-based infrastructure, such as TPM [160], is available to support secure booting, BinAuth can enjoy an increased security. The `Hashing_key` can also be stored securely by TPM.

5.5 Experimental Results and Discussion

We briefly discuss the impact of BinAuth on the system's performance. The two factors which impact the overall system performance are: (i) the Verifier's checking upon binary loading (invocation); and (ii) file modification monitoring. These factors affect the user's waiting time for process execution and file operations. We use both micro and macro benchmarks to determine the worst case and the average performance overheads.

The benchmarks are run on an Intel Core 2 Duo PC with 2GB of RAM running Windows XP with SP2. Each benchmark is run five times. As we want to investigate the effect of the cached Verifier, each benchmark is run with caching and without caching.

To see the difference of using different hashing algorithms, we implement and benchmark three algorithms: MD5, SHA-1 and SHA-256. For clarity, only the results of MD5 and SHA-256 are shown in Table 5.2. (The results of SHA-1 are always in between those of MD5 and SHA-256). When caching is enabled, results of different hashing algorithms are not distinguished (shown as Cached-MAC in the table), because the binaries do not require an MAC checking during the benchmark. This is since the first run is ignored[5], and no binary is modified during the benchmark.

To see the difference against digital signature based authentication systems, we also compare the performances of our scheme with those of the Microsoft official Authenticode utility called `SignTool` [102], and another Sysinternals (now acquired by Microsoft) Authenticode utility called `Sigcheck` [129]. Note that these two tools are user-mode programs. They are here to illustrate the difference between non-mandatory strategies used in Authenticode with our in-kernel mandatory authentication.

We have conducted a total of three benchmarks. The first two benchmarks investigate the system's performance. The third investigates the tradeoffs between the cached and the uncached verifications. The first two benchmarks investigate the system's performance under two following scenarios:

1. **Micro-benchmark:** The micro-benchmark aims to measure the worst case performance overhead incurred by the scheme. Note that this is primarily intended to measure the authentication cost, and not other overheads. Here, we have two micro-benchmark scenarios.

   (a) **Loading EXE:** This executes the `noop.exe` program, a dummy program that immediately exits, for 10,000 times. This scenario measures the overhead for authenticating the `EXE` file. We use different binary sizes (40 KB, 400 KB, 4 MB, and 40 MB respectively).

[5]When caching is enabled, we ignore the result of the first run because the authentication overhead is already shown in the uncached case. Even if we take the first run into account, its impact will be very small due to the high number of runs that the benchmarks are subject to.

Authentication System	Micro-Benchmark						Macro-Benchmark	
	`noop.exe` (40 KB)		`noop.exe` (40 MB)		`load-dll.exe`		`build`	
	time	slowdown	time	slowdown	time	slowdown	time	slowdown
Baseline	22.76	–	30.07	–	45.32	–	66.26	–
EXE Only:								
`SignTool`	2,822	11,637%	4,850	16,033%	73.49	62.16%	97.00	46.39%
`Sigcheck`	1,720	7,457%	5,629	18,623%	62.82	38.62%	110.5	66.72%
Uncached-MD5	25.96	14.08%	2,150	7,052%	45.34	0.05%	70.85	6.93%
Uncached-SHA256	30.29	33.07%	9,005	29,851%	45.34	0.05%	71.79	8.35%
Cached-MAC	23.20	**1.93**%	30.63	**1.88**%	45.33	**0.02**%	67.62	**2.06**%
All Binaries:								
`SignTool`	11,867	52,043%	14,030	46,565%	16,018	35,244%	–	–
`Sigcheck`	4,283	18,772%	6,186	20,478%	12,548	27,587%	–	–
Uncached-MD5	26.10	14.67%	3,881	12,811%	128.8	184.1%	79.31	19.69%
Uncached-SHA256	30.42	33.67%	9,302	30,839%	201.3	344.0%	91.80	38.55%
Cached-MAC	23.25	**2.14**%	30.58	**1.72**%	45.35	**0.07**%	67.88	**2.45**%

Table 5.2: System's performance benchmark results showing times (in seconds) and slowdown factors. The worst slowdown factor for each scenario is shown with underline, whereas the best is in bold. We define $slowdown_x = (time_x - time_{clean})/time_{clean}$.

(b) **Loading DLL:** This executes the `load-dll.exe` program for 100 times. This scenario is used to find out how loading DLLs impacts the system's performance. The program `load-dll.exe` loads 278 standard Microsoft DLLs with a total file size of ∼75MB. The size of the `load-dll.exe` itself is 60KB.

2. **Macro-benchmark:** The macro-benchmark measures the overhead under a user's typical usage scenario. Our benchmark is to create the Windows DDK sample projects using the `build` command. In each test run, 482 C/C++ source files in 43 projects are built. This benchmark is chosen as it is deterministic and non-interactive; and it creates many processes and uses many files.

We benchmark `SignTool` and `Sigcheck` in the following fashion. We first sign `noop.exe` and `load-dll.exe` using `SignTool`'s signing operation. We then measure the total times for authenticating and executing these two sample programs. For the macro-benchmark, we replace each development tool in the DDK (i.e. `build.exe`, `nmake.exe`, `cl.exe` and `link.exe`) with a wrapper program which first authenticates the actual development tool and then invokes it. For the micro-benchmark, we consider two settings: (i) `EXE` only; and (ii) all binaries (`EXE` + `DLL`). The macro-benchmark, however, only tests the `EXE`s only. This is because, during the macro-benchmark, many programs are invoked, and for each program each invocation may dynamically load a different set of DLLs. Thus, it is hard to keep track of what DLLs are loaded, and it is unfair to test the system with all the DLLs used.

The results are presented in Table 5.2. For clarity, we do not show the results for "`noop.exe` 400 KB" and "`noop.exe` 4 MB" because they are bounded by those

of "`noop.exe` 40 KB" and "`noop.exe` 40 MB" in a linear fashion.

We can see that the overheads of `SignTool` and `Sigcheck` makes them unusable if DLLs are to be checked (352 times slower on `load-dll.exe` case). If only `EXE` are checked, the overhead is at least ∼40%. Furthermore, using these two tools would also incur additional overheads from creating a process. Our main purpose here is just to show the difference between what can be done in user-mode versus in-kernel scheme. We can see that the uncached-MD5 and SHA256 are considerably faster than `SignTool` and `Sigcheck` in almost all cases. The only exception is the micro-benchmark case of loading huge 40MB `noop.exe`, where the uncached SHA256 runs slower than both `SignTool` and `Sigcheck`. This is since we have BinAuth repeatedly authenticate the 40MB-length binary for 10,000 times in each run using the more computationally expensive SHA256 operation. As mentioned earlier, the results are not fully comparable for this case since `SignTool` for example uses SHA-1 algorithm in creating a file's signature.

Now, let us compare the results of the cached and the uncached BinAuth. We note that the running times for the micro-bencmarks for the uncached is very high for SHA-256 on `noop.exe` 40MB. Recall that the binary is authenticated 10,000 in each run. The micro-benchmark results on the smaller file size shows that the uncached's overhead is acceptable (∼14–33%). The macro-benchmark for the uncached BinAuth however shows that under a typical usage, the uncached has overheads of ∼8%, while the cached brings this down to very small value of ∼2%. The cached's overhead is even almost neglible in the `load-dll.exe` benchmark (0.02%). Note that as the uncached overhead is quite small, the results are dominated by the non-determinism in timing measurements.

Moving to all binaries (`EXE` + `DLL`), we can see the effect of programs using many DLLs in Windows. (Typically, there are more codes contained in invoked `DLL`s than `EXE`.) The overhead incurred by caching for `noop.exe` 4KB is still small, while the uncached can grow to between 20-40% depending on the hash algorithm.

The third micro-benchmark investigates the tradeoffs between the cached and the uncached verification. The results are shown in Table 5.3. Using caching means that the MAC verification is amortized over executions, but it has added an overhead for monitoring file modifications. The uncached mode is just the opposite. Our micro-benchmark opens a file for writing 100,000 times to measure the worst case overhead incurred by the file modification monitoring. We have 3 experiments: (i) a clean system without binary authentication as the baseline case; (ii) cached BinAuth with a modified (protected) binary file; and (iii) cached BinAuth with a modified non-binary file.

The results for the file modification micro-benchmark show that for the binary authentication using caching, it does not matter whether the file being written to is a binary or not. File monitoring in both cases incur ∼60% overhead compared to a clean system. The uncached BinAuth incurs no overhead associated with monitoring potential

Case	time	slowdown
Baseline	4.051	–
Cached-MAC with a modified (protected) binary file	6.809	68.1%
Cached-MAC with a modified non-binary file	6.266	54.7%

Table 5.3: Benchmark results showing file modification monitoring overheads.

file modifications. This means that under some usage scenarios where files are frequently modified, the uncached strategy may be preferable over the cached one even when the Verifier's overhead is higher in the former.

5.6 BinAuth and Software_ID Scheme

We complement binary authentication with a scheme called Software_ID in order to simplify binary management issues. The idea is that we associate a *unique string* to a particular binary of a software product. This is similar to the recent CPE initiative by MITRE [104]. Software_ID is however different from CPE in that Software_ID is associated not only to a family of software, but with *each* individual (binary) file. Additionally, we suggest that Software_ID is to be signed together with a binary file. The benefits of combining binary authentication and Software_ID are twofold. Firstly, we can obtain the ID of an executed binary in an accurate manner right to *the file level*. This allow us to perform any necessary check, such as pre-execution vulnerability check (described in Chapter 6), on a binary. The vulnerability alert advisory may already contain the affected Software_ID. Compared to CPE, it is now possible to identify and block only the affected file(s) in a vulnerable software product.[6] Secondly, due to the use of binary authentication, a Software_ID string is securely protected against illegal modifications.

Software_ID should ideally come from the software developer. Alternatively, it can be assigned by the system administrator. The key to ensuring the uniqueness of Software_ID lies on its standardized format. We can define Software_ID as follows:

$$\text{Software\_ID} ::= \langle\ \text{vendor\_ID} \,||\, \text{product\_ID} \,||\, \text{module\_ID} \,||\, \text{version\_ID}\rangle.$$

Here, '||' denotes string concatenation. Module_ID records the file name of the binary. Different from CPE, version_ID does not keep track of a software product version. Rather, it is meant to keep track of *file versioning* in the case where a binary is updated or patched.[7] For the software product version information, we make it as part of product_ID. Similar to CPE, we can leverage on the possible existing trust infrastructures [165], such

[6]This policy of blocking only the affected file(s) can be acceptable, for instance, when the vulnerable components are the non-essential components of the software package.

[7]To deal with module updates, module_ID can also be used to record the module's version information. Having a separate version_ID, however, is useful to easily track different versions (or patched versions) of the same program.

as the domain name of the developer, for the vendor_ID in the absence of a centralized naming authority for software vendors.

Software_ID can also be used to help ensure controlled update on (protected) binaries. We can stipulate that a signed binary can only be replaced with another signed binary which satisfies the following requirements:

- The new binary is of the same file name.
- It is correctly signed by the same developer.
- The binary has the same Software_ID, but with higher version_ID information.

Upon a valid binary update, the database (`Digest_file`) is thus accordingly updated. Using this mechanism, we can therefore protect the binaries against old-attacks attempting to exploit the software update process. Note that this mechanism does not need an additional file/signature revocation list as in [8, 28].

5.7 Summary

We have analyzed the problem of providing program integrity assurance. We also have developed a framework which characterizes a binary authentication system, and how these design options may affect on ensuring software integrity on a host. Using BinAuth, we have shown how a mandatory in-kernel authentication system can provide an increased trust on good software execution with acceptable performance overheads even on a complex OS like Windows. BinAuth integrates well with PKI without having to heavily rely on it. Additionally, we have shown how BinAuth can be combined with a simple Software_ID scheme or CPE to simplify binary version management and vulnerability alert processing.

Chapter 6

Towards Automated Vulnerability Alert Processing

The number of security vulnerabilities discovered in computer systems has increased explosively [29]. In order to keep track of security alerts, administrators generally rely on vulnerability alert repositories/databases[1] such as [161, 135, 158]. Such databases are however designed primarily to be read and understood by humans. Given the speed at which an exploit becomes available once a vulnerability is known, and the frequency of occurrence of such vulnerabilities, manual human intervention becomes too slow, time-consuming and possibly ineffective [30] (see also background information in Section 2.4).

In this chapter, we address the challenge to provide a vulnerability-free execution of a program with respect to the (publicly) known vulnerabilities. This constitutes an important step in securing the PPLC. Given the limitations of the manual processing of alert information, we incorporate a framework for automating vulnerability alert processing on a host. The framework is based on our work [148], which was developed when vulnerability alert database for automated processing was still not addressed as widely as today. More specifically, our framework aims to achieve the following objectives:

- To define a coordinated vulnerability database, with entries that are meant to be machine-readable and processable. The database must be comprehensive and up-to-date in its contents, and is coordinated from multiple different existing sources in a well-structured manner. In our approach in [148], we assume a representation using the relational database model.
- To represent each alert information as a *specification data* using a vulnerability description expression scheme rather than a code (i.e. binary or script). The specification should enable efficient operation of a vulnerability scanner. In addition, it

[1] As previously mentioned, we follow a common practice to refer to a vulnerability alert repository as a vulnerability database. This is despite the alert entries are narrative and not structured using any well-defined database schema, e.g. the relational model.

should make it possible for the scanner to *relate* different vulnerability entries and determine the possibility of a "*chain of exploit*".

- To define and develop a proof-of-concept vulnerability scanner, which performs an automated alert processing and vulnerability scanning on a host. Such a scanner must be *general purpose* in that its mechanisms are independent from any specific entry definitions and not tied down to limited vendor-specific entries. Furthermore, the scanner operations must be minimal and should be *open* to possible scrutiny and verification.

Despite being published several years back, our results are still relevant with present efforts to secure a host system. Our proposed mechanisms in [148] serve as example mechanisms to achieve the above-mentioned objectives. Since the creation of an integrated database should involve many relevant parties, the mechanisms are however not meant to be a standalone definitive solution. Rather, the main contributions of our work [148] are the definition of the framework and the identification of its key components as well as properties required for an effective automated vulnerability processing. In fact, our proposed framework is in line with the on-going standardization efforts on automating vulnerability processing, such as OVAL [106], CPE [104] and SCAP [115]. With respect to these standardization results, we are pleased to see how the vulnerability management field has been progressing in recent years, and that our published work may have played a part in spurring the developments of machine-oriented vulnerability alert processing. We contrast later our framework and these standardized results. Moreover, we also point out how the standardization results still need to be extended to realize the fully-automated vulnerability processing solution as envisioned in our framework.

The rest of this chapter is organized as follows. Section 6.1 describes the existing works. The overview of our framework is given in Section 6.2. Our proof-of-concept database is presented in Section 6.3. Section 6.4 covers the corresponding vulnerability description expression scheme. A prototype scanner is described in Section 6.5. Section 6.6 discusses deployment issues as well as other aspects. Finally, Section 6.7 gives a summary.

6.1 Existing Works and Challenges

We discuss below various works related to the three components of the proposed framework, namely machine-oriented database, automated vulnerability scanner and vulnerability description scheme. We additionally highlight the challenges found in some of the systems, which have provided motivations for our framework's design goals.

6.1.1 Machine-Oriented Vulnerability Database

There exist a number of vulnerability databases which reorganize and integrate various vulnerability alerts into one repository that is searchable based on specified attribute values. Examples are ICAT[2], Public Cooperative Vulnerability Database [126], and the Open Source Vulnerability Database (OSVDB) [158]. Although these databases are searchable, they are however not specifically designed for any automated applications. Krsul [79] proposed a comprehensive taxonomy of vulnerabilities for possible further processing or automated manipulation. A database definition was also proposed. However, no specific applications were co-designed, or shown to work together, with it.

National Vulnerability Database (NVD) [157] is an active vulnerability database managed by NIST, based on a recent initiative supported by the U.S. Department of Homeland Security. It hosts the U.S. government repository of vulnerability management data, and can be searched based on CVE [105] or OVAL [106] query. NVD was specially created with the intent of reorganizing and aggregating vulnerability information from multiple existing databases into a standardized database. MITRE also maintains OVAL Repository [108], which hosts *OVAL Definitions*, i.e. machine-readable tests written in the OVAL Language [106]. These two databases are thus in line with our proposal for an integrated and unified machine-oriented database.

6.1.2 Host-based Vulnerability Scanner

There have been a number of popular tools that scan for any presence of vulnerability or configuration weaknesses in a system. Some examples are COPS [43], Ferret [137], SATAN [134], and Nessus [117]. In code-based scanners such as [43, 134], the logic of vulnerability checking is embedded tightly in the scanner's code. This means that including a new vulnerability check requires one to update the scanner's code or its subcomponent(s). Most recent scanners such as [137, 117] take a modular approach. A vulnerability check is usually represented as a "plug-in module", which can be fed to the scanner to perform the test. Nessus writes its vulnerability checks in a scripting language called NASL [10]. Vulnerability checks written in languages like NASL however can potentially be too powerful, which may allow for possible misuse. NASL resembles code rather than data. Thus, an attacker may target the plug-in modules so as to perform its own unauthorized operations on a target host. In contrast, our proposed framework uses a declarative language to specify a vulnerability check. The actions are performed by a generic scanner which is made available for third party inspection.

Windows Update [172] is a Microsoft online tool for automatically updating Windows OSes and and its installed components with recent patches. It illustrates some important

[2]The database used to be available at `http://icat.nist.gov/icat.cfm`. The site seems to be no longer in operational, and access to it is referred to the National Vulnerability Database (NVD) [157].

issues with vendor-specific automated tools. Windows Update and its more automatic cousin, Windows Software Update Services [171], are closed systems. They are examples of the "black-box update model". This leads to the following issues:

1. **Trust and privacy issue**: As there is no open specification or possible inspection on the scanner, no complete trust can be put on the scanner. It is difficult to determine if the scanner performs the correct actions while preserving the local system security policy. Since the update system is typically hosted on the vendor's server, there is also no guarantee that any local sensitive information will not leak to an external source.
2. **Non-standard vulnerability checking issue**: Windows Update behaves more as a vendor's patch-updating mechanism rather than a standardized vulnerability entry checking. Thus, little feedback is given back to users in terms of a standardized vulnerability report information. This might be too limiting for an administrator who, for example, wants to ensure that his/her systems are up-to-date against recent vulnerability reports regardless of whether patches for the vulnerabilities are available or not.

Thus, in contrast to this black-box and vendor-specific update model approach, we propose an open system which can cater to heterogeneous environments.

6.1.3 Vulnerability Description

Besides describing a vulnerability entry, our proposed vulnerability description scheme also aims to allow the scanner to analyze a potential *chain of exploit* involving multiple vulnerabilities. More specifically, the specification must take into account of the effects of a vulnerability in a way that allows us to verify whether they can induce the necessary condition(s) for the exploitation of other vulnerabilities.

Modeling vulnerabilities and their interactions can be traced back to the Kuang system [16]. Baldwin proposed a rule-based system called Kuang, and developed a sample Unix scanner called U-kuang [16]. Based on a set of deduced access privileges on a host, the system tries to find any ways the host's access policy can be violated. COPS [43] incorporated the U-kuang system into its host-based scanner. NetKuang [176] extended the rule set in Kuang to allow it to also work in a networked environment. Based on the concepts from Kuang, the KuangPlus system [63, 159] was developed as a Perl-based scanner. Our prototype scanner is also written in Perl, and is inspired by KuangPlus due to Perl's versatile text processing facilities. Given this Perl-based scanner, we propose our vulnerability description scheme which we find easy to describe and amenable for vulnerability-chaining analysis. Unlike KuangPlus [63, 159], we take a vulnerability check as a declarative data rather than a code.

Open Vulnerability and Assessment Language (OVAL) [106] is a security assessment language from MITRE, which is supported by the U.S. Department of Homeland Security. The OVAL Language specifies how to check a host for the presence of software vulnerabilities using XML. The vulnerability checks are to be verified by a scanner called OVAL Interpreter. The present OVAL Definition Language however does not allow us to abstract the effects of the exploitation of a vulnerability. [92, 33] presented the DESEREC vulnerability definition language, which extends our Movtraq description language [148] by using the OVAL format in order to enable vulnerability-chaining analysis.

6.2 Movtraq Framework: System Overview

We call our proposed framework for automated host vulnerability alert processing as *Movtraq (Machine Oriented Vulnerability and Tracking)*. Figure 6.1 shows the system overview of Movtraq. The integrated Movtraq vulnerability database is designed to be compiled from multiple sources and is usable directly by an automatic scanner.

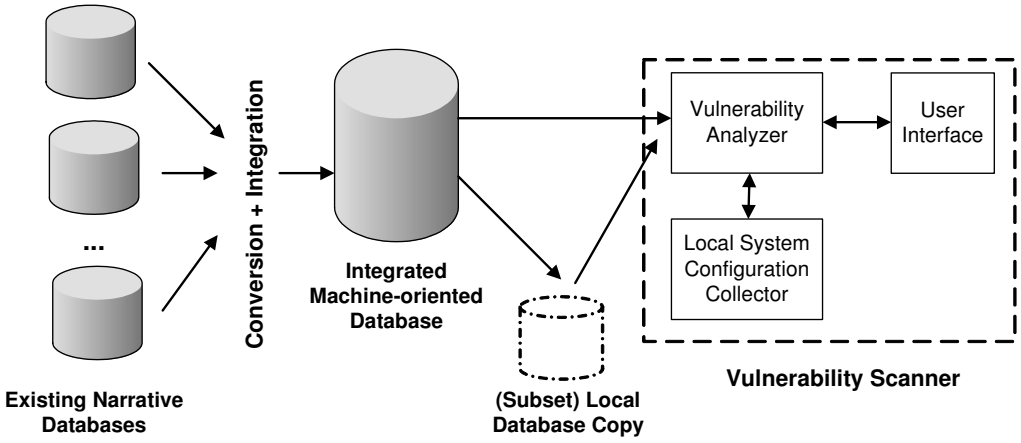


Figure 6.1: System overview of Movtraq automated framework, showing the vulnerability database and scanner. Note that (subset of) the database may be replicated in the target's host or a proxy server within the same administrative domain.

We now explain the three main components of the framework, namely: integrated database scheme representation, vulnerability description expressions, and automated sample scanner.

6.3 Movtraq Vulnerability Database

6.3.1 Design Goals

As mentioned, our philosophy for Movtraq database is that vulnerabilities should be stored as well-structured data (description). Hence the data can be stored in a database (or any data description language, e.g. XML). Our database is designed with the following criteria:

- Each vulnerability entry includes the general information of the vulnerability, the conditions (*pre-requisites*) in which it can affect a host, and its *consequences*.
- The pre-requisites and the consequences of a vulnerability are described using an abstraction which we call a *vulnerability description expression scheme*. It allows a precise formulation of the nature of the vulnerability and is machine processable.
- The structure of the database should allow easy retrieval both by users and automated-tools via SQL.

6.3.2 Content of a Vulnerability Entry

The main challenge in designing the new database is to determine what the actual contents of each vulnerability entry should be. For our proof of concept, we focus on what the database *should contain* rather than on constructing an optimal database schema. The data fields corresponding to a vulnerability entry fall into the following three categories:

General Information Fields

This portion mostly contains references to several public vulnerability databases such as CERT, Bugtraq, etc. The purpose of these fields is to give the user a reference to the original source of information. This is mainly for human understanding.

Vulnerability Conditions

This second category provides the main content of the machine processable vulnerability information. A vulnerability normally exists within a context. Therefore, it is described in terms of its *component factor* and associated *environmental factors*. By "component factor", we mean the system component, i.e. application or OS, where the vulnerability originates. "Environmental factors" refers to the settings or services in the local host which make it subject to the vulnerability.

We distinguish vulnerabilities into two types:

- vulnerability which currently exists on the system; and
- vulnerability which *potentially* exists on the system.

There are four different combinations of checking results on component and environmental factors, which are diagrammatically shown in Table 6.1:

		Environmental Factor	
		Match	**No-Match**
Component Factor	**Match**	*Vulnerability Exists*	*(Potential) Vulnerability*
	No-Match	*Vulnerability Free*	*Vulnerability Free*

Table 6.1: Combination of checking results between component and environment factors.

Case 1: Component factors: *match* & Environment factors: *match*. We will get this result when a particular vulnerability's component exists on the local system and the settings of local system match all the required environment factors. In this case, we will conclude that *the vulnerability exists* on the system.

Case 2: Component factors: *match* & Environment factors: *no match*. This occurs when we can detect a vulnerable component on the local system, however the settings of the local system does not match the environmental requirements. While the vulnerability is currently not applicable yet, it has the *potential* to affect the system should its environment change. For example, consider the case of "Apache Web Server Chunk Handling Vulnerability" (`http://www.cert.org/advisories/CA-2002-17.html`). Even if Apache is installed in our system, we will not be affected by this vulnerability as long as we do not provide http services.

Case 3: Component factors: *no match* & Environment factors: *match*. In this case, the vulnerability would appear to be not applicable. However, there is a subtle issue. Consider the case of OpenSSL which previously had several exploitable stack overflow vulnerabilities (`http://cve.mitre.org/cgi-bin/cvename.cgi?name=CAN-2002-0656`). OpenSSL may not be installed as an individual component, so even if there is a database entry for the OpenSSL vulnerability, this would return a negative result in terms of the vulnerability component factors. However, OpenSSL is commonly included in applications such as Apache, Sendmail and Bind. Thus, it is necessary to check for the existence of such applications. To deal with this case, we therefore need the vulnerability entry to list all the affected applications based on the dependency of the component factors.

Case 4: Component factors: *no match* & Environment factors: *no match*. The vulnerability does not exist on the local host. However, the remark on vulnerability component dependency in Case 3 is also applicable here.

Vulnerability Impact (Consequences)

The third category of data concerns the impact of vulnerability, which describes the possible consequences of a vulnerability if it is successfully exploited. In our database, it is represented using the vulnerability description expression. This enables checking of the relationship among different vulnerabilities and whether they can affect one another.

6.3.3 Database Structure

As we have argued, the exact structure of the database is not very important. Rather, it is the content and having it in a machine processable format. In our proof-of-concept design, the database has seven main entities (tables). The description of these entities is given in the Appendix B.

6.4 Vulnerability Description Expressions

As shown above, the (environmental) pre-requisites and consequences of a vulnerability entry require a machine friendly specification. After studying ∼1,000 vulnerability alerts from CERT advisory database, we found that most of the information for these two categories can be described effectively using the *vulnerability description expressions* as describe below. These expressions are designed to be easily processable using text-processing based vulnerability scanners, such as our Perl-based prototype scanner inspired by KuangPlus [63, 159].

An expression is written with the following syntax:

$$\langle Vulnerability\_Expression \rangle := \langle Target\_Object \rangle \mid \langle Action \rangle \, \langle Target\_Object \rangle. \quad (6.1)$$

An action is written prefixed by '@'. Table 6.2 illustrates the actions and the types of the corresponding target objects. Here, we have used a concise notation mainly for brevity. A real system may also support multiple syntaxes.

Rather than giving a formal definition of the target objects, we have listed the target object examples in Table 6.3. The following prefixes are used: '%' is used to denote an actual value; '#' is used to denote a symbolic value; and '&' is used for expressing users/groups associated with an application/service.

As our proof-of-concept implementation is for Unix systems, the examples and objects are also Unix based. Vulnerabilities for other OSes may require extension to the types of target objects and actions.

Syntax	Semantics
@G $\langle u \vert g \rangle$	Gain $\langle$ *user_object* u | *group_object* $g \rangle$
@R | W $\langle f \vert m \rangle$	Read | Write $\langle$ *file_object* f | *memory_object* $m \rangle$
@A $\langle f \vert m \rangle$	Access (read and write) $\langle$ *file_object* f | *memory_object* $m \rangle$
@C $\langle f \rangle$	Create $\langle$ *file_object* $f \rangle$
@K $\langle f \vert m \rangle$	Corrupt $\langle$ *file_object* f | *memory_object* $m \rangle$
@X $\langle f \vert c \rangle$	Execute $\langle$ *file_object* f | *code_object* $c \rangle$
@S | I $\langle n \vert a \rangle$	Crash | Disrupt $\langle$ *node_object* n | *application_object* $a \rangle$
@D $\langle n \vert a \vert s \rangle$	Deny $\langle$ *node_object* n | *application_object* a | *service_object* $s \rangle$
@U $\langle r \rangle$	Use $\langle$*resource_object* $r \rangle$
@E $\langle r \rangle$	Exhaust $\langle$*resource_object* $r \rangle$

Table 6.2: Actions in the Vulnerability Description Expressions.

6.4.1 Examples using Vulnerability Expressions

The following examples use the expressions to describe various vulnerability consequences.[3]

- @D n#N: Denial of Service for the whole network (Ref: Cisco IOS Interface Blocked by IPv4 Packet CERT_ID VU#411332).
- @G u#S : Gain superuser right (Ref: Linux Kernel Privileged Process Hijacking Vulnerability, Bugtraq_ID: 7112).
- @G u#R : Gain remote user right (Ref: Apache htpasswd Password Entropy Weakness, Bugtraq_ID: 8707).
- @R f%/etc/passwd : Read file `/etc/passwd`.
- @X f#*(4777) : Execute a file with setuid permission.

The following are examples of portions of the machine oriented fields in the database for several vulnerabilities:

- *MySQL Password Handler Buffer Overflow Vulnerability*:
  CVE_ID: CAN-2003-0780
  Bugtraq_ID: 8590
  Vul_Con: @G u%mysql; @G u#L; @X c#(u%mysql)
  Vul_OS: null
  Vul_App: (*Various Mysql versions*)
  Env_User: u#L
  Env_File: null
  Env_Remote: No
  Exploit: No
  Env_OS: null
  Env_App: mysql

[3]For simplicity, multiple expressions are separated by semicolon.

Syntax	Semantics
⟨**u**⟩:	**User_Objects:**
u#R	Remote user
u#L	Local user
u#S	Super user
u#*	All users
u#P	Physical user
u#U	User whose privilege is beyond that of current user
u%⟨*uid*⟩	User with UID=*uid* (e.g. 100)
u%⟨*user_name*⟩	User with specified *user_name* (e.g. 'nobody')
u&App	User running corresponding application process
u&Svc	User running corresponding service (i.e. daemon)
u&Kernel	User who can access or control OS kernel
⟨**g**⟩:	**Group_Objects:**
g#*	All groups
g&⟨*App*|*Svc*⟩	Group of corresponding application process/service
g%⟨*gid*⟩	Group with GID=*gid* (e.g. 50)
g%⟨*group_name*⟩	Group with specified *group_name* (e.g. '`sys`')
⟨**f**⟩:	**File_Objects:**
f#*	All files
f#passwd	Pathname corresponding to passwd file
f#shell	Pathname corresponding to shell files, e.g. `/bin/bash`
f#system	Pathname corresponding to system files in OS
f#*(4777)	All files with permission 4777
f#F	Files beyond current user access rights
f%⟨*file_name*⟩	File *file_name* (e.g. '`/etc/passwd`')
f&App	File associated to running application process
⟨**m**⟩:	**Memory_Object:**
m#M	Memory area beyond current user's access right
⟨**n**⟩:	**Node_Objects:**
n#S	Target node where an application is installed and/or related service is running
n#L	Nodes in local area network
n#N	Network
n%⟨$IP_1[-IP_2]$⟩	Node with specified IP address (may be a range)
⟨**a**|**s**⟩:	**Application_Objects and Service_Objects:**
a%⟨AppName⟩	Application name (e.g. as listed by '`ps`' command)
s%⟨SvcName⟩	Service name
⟨**r**⟩:	**Resource_Objects:**
r#M	Memory
r#CPU	CPU
r#B	Network bandwidth
r#D	Disk space
⟨**c**⟩:	**Code_Object:**
c#(⟨u⟩)	Piece of code with execution privilege of user_object u, e.g. privilege escalation

Table 6.3: Objects in the Vulnerability Description Expressions. The prefix '%' is used to denote an actual value, '#' for a symbolic value, and '&' for expressing users/groups of an application/service.

- *Linux Kernel IOPERM System Call IO Port Access Vulnerability*:
  CVE_ID: CAN-2003-0246
  Bugtraq_ID: 7600
  Vul_Con: @A f#F
  Vul_OS: (*Various Linux distributions*)
  Vul_App: null
  Env_User: u#L
  Env_File: null
  Env_Remote: No
  Exploit: No
  Env_OS: Linux kernel 2.4.0 - 2.4.21, 2.5.0 - 2.5.69
  Env_App: null

- *Linux 2.4 Kernel execve Race Condition Vulnerability*:
  CVE_ID: CAN-2003-0462
  Bugtraq_ID: 8042
  Vul_Con: @A f#F; @X c#(u#S); @G u#S
  Vul_OS: (*Various Linux distributions*)
  Vul_App: null
  Env_User: u#L
  Env_File: f#*(4111)
  Env_Remote: No
  Exploit: Yes
  Env_OS: Linux Kernel 2.4.0 - 2.4.21
  Env_App: null

- *Multiple Vulnerabilities In OpenSSL*:
  CVE_ID: CAN-2002-0656
  Bugtraq_ID: 5363
  Vul_Con: @X c#(u&App); @G u#L
  Vul_OS: null
  Vul_App: (*Various Apache versions and OpenSSL-based applications*)
  Env_User: u#R
  Env_File: null
  Env_Remote: Yes
  Exploit: Yes
  Env_OS: null
  Env_Svc: Corresponding service provided by the vulnerable application

6.4.2 Translation Issues

From our experiments in translating text-based vulnerabilities into vulnerability expressions, we encountered the following issues:

- The vulnerability description in the narrative database sources is sometimes rather vague. Some examples are: "could expose sensitive information to local attackers" (Bugtraq_ID 8233), "gain access to sensitive information" (Bugtraq_ID 9558), or "leads to unauthorized access to attacker-specified resources" (Bugtraq_ID 9778). We however require a more precise consequence, which either means settling for a general (high-level) consequence or that much more work is required to understand the vulnerability.
- Our vulnerability expression language is designed to capture general expressions at the OS level. It does not express various application specific descriptions, such as: "to access variables outside the Safe compartment" (Perl, Bugtraq_ID 6111), or "could compromise the private keys of ElGamal signing key implementation" (GnuPG, Bugtraq_ID 9115). Such consequences are approximated by our translation into the closest vulnerability expressions capturable by our language. In the two examples above, we can rewrite them into: access of memory and files beyond the current user's right, respectively.
- Some vulnerability consequence entries, particularly those of CAN (didate) type, are listed as "unknown consequence" (e.g. Bugtraq_ID 10428). Hence, we either have to ignore such consequences until they are known and specified, or use a special form to indicate unknown consequences.

6.5 Movtraq Vulnerability Scanner

6.5.1 Design Goals

Our objective is build an automatic scanner which can use the defined database to perform the following tasks:

- Check whether a given vulnerability exists on a local system;
- Scan the local system for all possible vulnerabilities;
- Notify the existence of potential vulnerabilities on the local system;
- Analyze the relationship among different vulnerabilities, e.g. whether one vulnerability can be exploited to lead to another.

6.5.2 Implementation

To demonstrate the use of the Movtraq database, we implement a prototype of automatic vulnerability scanner called the *Movtraq vulnerability scanner*. The scanner runs on two different versions of Unix, namely Red Hat Linux and FreeBSD. This is to demonstrate some degree of platform independence.

Our prototype scanner runs as a user-mode application following its invocation by the administrator. An alternative operation mode is to integrate it into the kernel, which then checks a particular binary right before the binary's execution. To avoid any significant delay in binary execution, this pre-execution vulnerability scanning would be more feasible if the host maintains a local (up-to-date) copy of the vulnerability database, or that the database is hosted by a server within its network. These possible deployment scenarios of the vulnerability database are discussed more in Section 6.6.1. To reduce the database size, the host can choose to copy only the relevant entries based on its installed software packages. In our prototype, which is aimed at highlighting the feasibility of the machine-processable database and corresponding automated scanner, we opt to implement a user-mode scanner which can be run interactively by the administrator.

The overall structure of the scanner together with the database is depicted in Figure 6.1. The integrated Movtraq database is stored in MySQL. The scanner consists of a local system configuration collector which collects information about the OS (which processes are running, which ports are open, hardware details, etc.), applications, and services on the system. Software versions are obtained by using the `rpm` utility on Red Hat and the `pkg_info` utility on FreeBSD. The scanner is written in Perl, and queries the MySQL Movtraq database using SQL.

An abbreviated sample log from running the scanner illustrates how OS, application, and environmental checkings are performed:

```
1. Apache Mod_Auth_Any Remote Command Execution Vulnerability
   Application version check: positive.
   Application environment (service port) check: negative.
   Conclusion: source application is detected, default port required
   is not open. Potential vulnerability exists but does not affect
   current system configuration.

2. Sun One/iPlanet Web Server Vulnerability to DOS
   Application version check: negative.
   Conclusion: source application not detected, safe from vulnerability.

3. Linux Kernel IOPERM System Call IO Port Access Vulnerability
   OS version check: positive.
   OS environment (running kernel version) check: positive.
   Conclusion: vulnerability detected!

4. MySQL Password Handler Buffer Overflow Vulnerability
   Application version check: positive.
   Application environment (process object) check: positive.
   Conclusion: vulnerability detected!
```

Only some of the pertinent checks from the log are shown above to illustrate the following points:

Example 1: Apache vulnerability exists, but the environmental factor check fails since the required port is not open.

Example 2: no vulnerability since the vulnerable application is not installed.

Example 3: OS vulnerability exists, so only OS component and the (environmental) kernel status checking are used.

Example 4: vulnerability inherent to MySQL, OS environment checking is skipped as it is not required.

6.5.3 Vulnerability-Chain Analysis

An interesting use of the scanner is that it can be used to test if existing vulnerabilities can be combined (chained) together to create more serious vulnerabilities. This mimics what a hacker might do to take advantage of indirect weaknesses on the system.

Consider the following example which is typical of a privilege escalation attack. Suppose the system has the following two vulnerabilities:

Name:	Buffer Management Vulnerability in OpenSSH	
Vul_ID:	57	
CVE_ID:	CAN-2003-0693	Bugtraq_ID: 8628
Vul_Con:	@G u#L	
Vul_OS:	null	Vul_App: Openssh apps
Env_Usr:	u#R	Env_File: null
Env_Remote:	Yes	Exploit: No
Env_OS:	null	
Env_App:	(Service provided by the vulnerable applications)	

Name:	Linux 2.4 Kernel execve Race Condition Vulnerability	
Vul_ID:	48	
CVE_ID:	CAN-2003-0462	Bugtraq_ID: 8042
Vul_Con:	@G u#S	
Vul_OS:	Linux	Vul_App: null
Env_Usr:	u#L	Env_File: f#*(4111)
Env_Remote:	No	Exploit: Yes
Env_OS:	Linux kernel 2.4.0 - 2.4.21	
Env_App:	null	

In this example, the scanner discovers that both vulnerabilities 48 and 57 are present. From Vul_ID: 57, a remote user (u#R) can gain the local user access (@G u#L), and this can chain onto Vul_ID: 48 which has two local environment requirements (i.e. the local user access: u#L and a setuid executable file:f#*(4111)). Thus, the scanner discovers that a remote user may be able to exploit the two vulnerabilities to gain the local root access.

The vulnerability-chaining analysis illustrates the benefit of a machine oriented approach and the use of vulnerability expressions to analyze the relationships among the vulnerabilities.

6.6 Discussion

6.6.1 Deployment Strategies for Movtraq

The prototype Movtraq system is sufficiently useful to be deployed in a number of ways. Some of the potential scenarios are depicted in Figure 6.6.1:

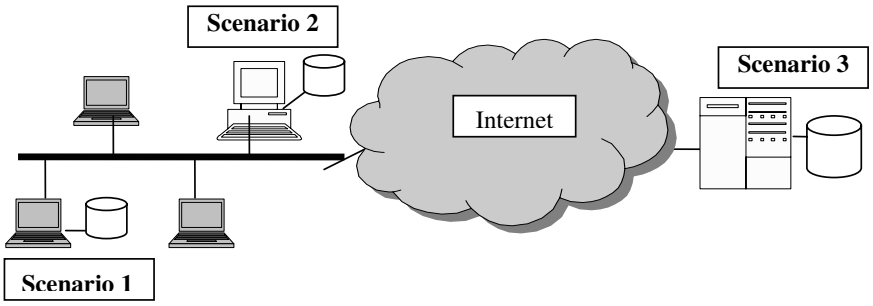


Figure 6.2: Deployment options for Movtraq vulnerability database.

Scenario 1: Local vulnerability database, local scanner.

Here, each local machine hosts its own database. The Movtraq database is meant to have been downloaded (securely) from an authoritative server. This scenario has the advantage that all operations can be done locally. The disadvantage is that an up-to-date database has to be maintained on every host.

Scenario 2: Organization-wide database, local scanner.

This simply extends Scenario 1 to an organizational context where there is an organization-wide database server. Where multiple machines have exactly the same configuration, one may choose to check only a subset of the machines.

Scenario 3: Internet-based database, local scanner.

Lastly, as in the automated update systems, a database server somewhere on the Internet serves as the database repository.

6.6.2 Movtraq and Recent Standardization Efforts

Our Movtraq framework was proposed several years back when machine-oriented vulnerability database and automated processing were still not widely addressed. At present, there exist several on-going collaborative standardization efforts, including those conducted by the U.S. Government-sponsored organizations, such as OVAL [106], National Vulnerability Database (NVD) [157], CPE [104] and SCAP [115].

Both OVAL Repository [108] and NVD [157] were specially created to be a unified and machine-oriented vulnerability database. These two databases are thus in line with our proposal of Movtraq vulnerability database. The OVAL initiative also has a language, called OVAL Language [109], which specifies how to check a host for the presence of vulnerabilities. A vulnerability entry is specified in XML in order for a scanner, called OVAL Interpreter, to perform the corresponding check(s) on a target machine. A sample OVAL Interpreter, although not a fully functional scanning tool yet, is freely available [107]. Thus, as an on-going collaborative efforts, the OVAL initiative does share the same philosophy and many common goals with our proposed framework.

The present OVAL Definition Language however does not allow us to abstract the effects of the exploitation of a vulnerability. As such, the language does not facilitate easy support for the vulnerability-chaining analysis, which we aim to provide with Movtraq. This deficiency led [92, 33] to present their own definition language, which extends OVAL definition with ideas taken from our Movtraq vulnerability description expressions. They describe vulnerabilities in terms of the conditions (*pre-conditions*) that make them exploitable and the effects (*post-conditions*) on the victim system.

In our present approach with Movtraq, objects involved in the checking of environmental factors are stored as database fields. The scanner is assumed to implicitly know how to operate on them. For an actively-evolving standardization effort like OVAL, such a rather rigid model may be too limiting. We notice that OVAL still lacks a set of standardized basic checking operations that can be performed on a target host. A database model worth investigating is thus one which standardizes possible scanner's basic checks on the target machine. Such standardized checks could also be parameterized to make them as generic as possible. These checks are then recorded *as database entries*. An entry specifies how the check should be performed (on a particular applicable OS), fields containing the values of objects involved in the checking, and notational descriptions of the pre-conditions and the consequence as in Movtraq. The capability of a scanner are thus basically limited by these defined checks. This would answer the growing nature of on-going standardization effort on vulnerability language. At the same time, it will provide standardized operations, and allows for vulnerability-chaining analysis.

Additional possible further enhancements on Movtraq, which are also applicable to other standards mentioned above, could include: convenient GUIs, compatibility with various popular OSes, a fully featured implementation, and a more sophisticated vulnerability model.

6.7 Summary

We believe that there is a pressing need for vulnerability databases which allow for automated host vulnerability processing. We have demonstrated a proof-of-concept database which allows effective integration of vulnerability data from multiple sources and can be used directly by a scanner. The proposed Movtraq scheme decouples the vulnerability information from its processing by storing a vulnerability description as declarative data. The scanner simply collects a system's local information and analyzes whether vulnerabilities apply on that host. Due to our use of the vulnerability description scheme in describing a vulnerability's pre-requisites and consequences, the analysis of potential chain of exploit involving multiple vulnerabilities becomes possible. Our prototype system was designed for Unix-based systems. However, we believe that the fundamental concepts in our design should still be substantially applicable to other OSes.

Chapter 7

Lightweight and Near Real-Time Certificate Revocation Schemes

As mentioned earlier, timely and efficient certificate revocation service is a critical pre-requisite for host security. This revocation service must be reliably available to ensure that public keys belonging to external parties are not revoked. In our PPLC context, the service is thus essential for secure software distribution, including possible mobile code (e.g. applet) execution on a local host (Step 1 in PPLC); or public-key based interaction of a running program with external hosts (Step 4). Disregarding such revocation verification would result in a catastrophic compromise to a host security.

In X.509-based PKI [68], certificate revocation is mainly supported by Certificate Revocation List (CRL) [36]. However, CRL is widely perceived to be costly, and is attributed as one of the main impediments to a successful global PKI deployment [88, 58]. Among others, CRL suffers from high bandwidth requirement and (generally) low timeliness guarantee. Several alternatives have been proposed, such as Online Certificate Status Protocol (OCSP) [113]. OCSP offers a potentially real-time recency, but requires the Certificate Authority (CA) to respond to any incoming query in real time with a signed message. As such, it imposes significantly higher computational and network requirements on the CA. Additionally, there exists a privacy issue since the information of all verifiers validating a certificate becomes known to the CA.

In this chapter, we look at a satisfactory revocation solution to support two important emerging trends. Firstly, mobile devices on the Internet will turn out to be a dominant source of Internet transactions. Thus, any revocation solution should be lightweight in terms of computation, bandwidth and storage on the verifier. Secondly, prevalent malware threats will increase the need for timely revocation supports, eventually towards real-time timeliness guarantees. The lack of lightweight and timely revocation today means that many applications, particularly on bandwidth- and processing-limited clients, are unable to offer proper revocation checking [131, 88]. Although this situation is slowly

changing, it is still unsatisfactory given the above trends.

More specifically, our goal is to have a practical revocation solution which provides *near real-time* revocation timeliness guarantees (e.g. from 1 to 10 minutes) as well as fast verification and low bandwidth on the verifiers. In addition, we would like to avoid excessively high overheads on any of the entities. This is to avoid any undesirable service bottlenecks. Lastly, we also aim to provide privacy assurance so that queries from the verifiers are unknown to external parties, including the CA. Although the CA is trusted for managing certificates, transaction information involving a certificate *should not necessarily* be made available to the CA.

Our proposal takes advantage of the "Extended-Validation SSL Certificate" (EVC) initiative [27] by the CA/Browser Forum. We propose a family of revocation schemes based on a simple but powerful paradigm of utilizing EVC principal's server as the distribution point for its own Certificate Status Information (CSI).[1] Our schemes allows for the co-location of Web (transaction) server, where a principal's certificate is normally obtained by the verifiers, and the corresponding CSI source.

This (seemingly small) change has a number of important advantages. The CSI distribution servers are self-managed (i.e. independent from the CA) and have a strong incentive to provide reliable and timely CSI services for the respective principal's transactions. Thus, it addresses the incentive problem on CSI management and dissemination [58]. Scalability issue is also naturally addressed since CSI accesses are distributed to the relevant servers which would be expected to have sufficient capacity to meet the respective transaction volumes. We still have overheads associated with CSI transfer management between the CA (or its trusted proxy) and EVC principals with our schemes, but we show that the overheads are manageable.

As a verifier obtains the CSI from EVC principals' servers instead of a (trusted) repository, we thus need revocation schemes whose CSI carries the "*liveness*" status of each *individual* certificate rather than the aggregated black-listing approach as in CRL. Two existing revocation schemes, CRS/NOVOMODO [100, 101] and OCSP [113] meet our needs. We enhance them to derive two schemes (CREV-I to CREV-II) that fit into the EVC-based setting and provide better performances.

Another contribution of this chapter is the use of a realistic model and a cost analysis framework for evaluating certificate revocation schemes. We follow the approach taken in [91, 65] which uses empirical revocation data from VeriSign. Our framework is a substantial extension of [91, 65] by incorporating the followings: revision of the CRL size derivation using a more accurate calculation method; and generalization of the CRL size estimate with certificate generation and the CRL issuance in minute(s) instead of day(s).

[1] The term CSI [66] is used in this thesis to denote information pertinent to the validity of a certificate. Thus, it encompasses: CRL data, hash tokens in CRS/NOVOMODO, and OCSP messages.

Furthermore, our analysis incorporates more realistic features in order to derive more realistic derivations of costs, which include: a more realistic query model on revoked certificates, the derivation of query probability on revoked and valid certificates, and more accurate accounting of message sizes based on the standardized CSI formats.

The remainder of this chapter is organized as follows. Section 7.1 gives a brief survey of related revocation schemes and cost analysis works. Section 7.2 provides additional background on EV Certificate, CRS/NOVOMODO, and the cost analysis of [91, 65]. We explain our CREV schemes in Section 7.3. We then elaborate our realistic cost-calculation framework, and analyze CREV schemes as well as the existing ones in Section 7.4. We discuss the comparison results in Section 7.5. Section 7.6 finally concludes this chapter.

7.1 Terminology and Related Works

Throughout this chapter, we will refer to the subject of a certificate as *principal*, EVC principal as *EVCP*, and the party verifying certificates as *verifier*. We refer to *Certificate Management Ancillary Entity (CMAE)* as a generic term for repository or directory, from which a verifier may obtain the CSI.[2] We also follow the definition of *revocation timeliness guarantee* in [3] as the time interval between when a CA makes a revocation record and when it makes that information available to the verifiers. In a per-certificate revocation scheme, such as CRS/NOVOMODO, OCSP and ours here, we assume that the CA will not unnecessary delay publishing the CSI indicating a revocation event in contrast to a periodic batch-processing mechanism.[3]

We briefly summarize the related existing revocation schemes below. Our focus here is to highlight their main strengths and weaknesses, and also their differences with our schemes. Section 2.5 provides more detailed background information on some of the surveyed schemes.

Standardized in X.509 [68] and also profiled in the IETF [36], CRL is also the most widely supported revocation scheme. It is however widely known to have serious shortcomings. Firstly, the CRL may eventually grow to a cumbersome size in very large PKIs, e.g. ∼750KB in Verisign's [131] and ∼40MB in the U.S. DoD's [118]. Secondly, the downloaded CRLs may be mostly useless given that more than 90% of the information is irrelevant to the verifiers [131]. Lastly, CRL does not offer adequate timely revocation guarantees. It is common for a CA to update its CRL only *daily*, as suggested in [164], although the CA may have a service for a shorter window period presumably with a higher charge.

[2] A CMAE may be trusted or non-trusted depending on the revocation scheme.

[3] Although OCSP is considered as a real-time service, its timeliness is only as up-to-date as the latency involved in obtaining the CSI from the Responder's definitive source. In our work here, we take an optimistic assumption that a CA providing OCSP service would provide *the lowest latency it can afford* as permitted by its applicable policy, in order to take advantage of the available real-time setting.

There are several methods for improving the basic CRL mechanism, such as CRL Distribution Points, Delta CRLs, and Indirect CRLs (see [3] for a survey). However, all these schemes still put the same requirement on the verifier to obtain a complete revocation list, which includes the unrelated entries. Furthermore, for our setting of EVCP servers as CSI distribution points, we require a scheme that carries a certificate's liveness status instead of the black-listing approach taken by CRL-based mechanisms.

OCSP [113] was proposed to provide a more timely certificate status checking. An OCSP Responder is required to return the status information about a specific certificate in a digitally signed response. Since OCSP is an online service, it necessitate a prompt and reliable OCSP Responder system with a high level of security. As such, it imposes significantly higher computational and network requirements on the Responder [85] (see also Section 2.5). Additionally, there exists a privacy issue with OCSP. Since the Responder is consulted whenever a verifier validates a certificate, it therefore knows all verifiers dealing with a principal. In our proposed schemes, the verifier obtain the CSI of an EVC principal directly from the principal's server. Hence, our schemes does not reveal certificate status queries to any third party, including the Responder (CA).[4] Although the CA is trusted for managing certificates, the transaction information involving a principal and a verifier should not be made available to the CA.

Several modifications have been proposed on OCSP, such as H-OCSP [111] and MBS-OCSP [20]. Similar to our proposed scheme (CREV-I), H-OCSP and MBS-OCSP make use of a hash-chaining technique. However, they require the CA to cater for a *potentially large number of verifiers*. Hence, the CA's bandwidth and storage requirements remain high. Our CREV schemes operate differently since the CA maintains hash chains belonging to EVC principals, whose number is much smaller than that of the verifiers. Moreover, we also employ a session-based hash chain as to reduce costs further (see Section 7.3).

Aiello et al. [5] proposed an improvement to CRS called "Hierarchical Scheme" aimed at reducing the CA-to-CMAE communication while still maintaining a low query communication. The improvement however comes at the price of a significant increase in the certificate size. Kocher [75] proposed Certificate Revocation Tree (CRT) which employs Merkle Hash Tree (MHT). The scheme however suffers from a high computational cost needed to update the CRT. Naor and Nissim [114] subsequently extended the CRT by using a more suitable data structure, 2-3 tree. All these three schemes are however rather different from CRS/NOVOMODO since the employed data structure maintains the statuses of *all* revoked certificates in contrast to an individual certificate status.

One of our proposed schemes, CREV-II, can utilize an optimization technique where

[4] In this thesis, we assume a scenario of OCSP deployment where the OCSP Responder is co-located with the CA.

an OCSP Response carries the status of multiple certificates. The technique is based on the certificate range definition in CRT [75]. A similar technique was also proposed by Koga et al. [76] earlier. Our use of this technique in CREV-II thus demonstrates how it can be utilized in the assumed network setting. In addition, using the developed realistic analysis framework, we also show how well the optimization can reduce the overheads.

Our realistic model and performance analysis framework offer a more realistic cost calculation. Below, we briefly describe the differences of our developed framework from previous analysis works such as [177, 84, 91, 65].

The work by Zheng [177] is one of the most widely cited works on cost analysis of certificate revocation schemes. However, it simply assumes that a certificate is always revoked at the half of its issued lifetime. In addition, its analysis of NOVOMODO is based on an assumption that the hash token is released on a daily basis. In contrast, our framework is based on a realistic revocation model which enhances the one proposed in [91, 65]. We consider a revocation service with timeliness guarantee on the order of minute(s) instead of day(s). Moreover, in Section 7.4.2, we derive a realistic query model on revoked certificates. Given the timeliness guarantee in minute(s), this model has a significant effect on the cost calculation of hash-chaining based schemes. This is because there is a great difference in cost between verifying a valid and a revoked certificate.

Lim and Lakshminarayanan [84] also conducted a performance analysis of various revocation schemes. They proposed a more detailed framework than that of Zheng [177]. However, the work is still based on the same assumption that a certificate is revoked at the half of its lifetime. The work derived the probability that a queried certificate is valid/revoked. This allows for a more accurate cost estimate in hash-chaining based schemes. However, the probability derivation is simply based on *the number of valid and revoked (but not yet expired) certificates* in the certification system. In contrast, our work (see Section 7.4.2) develops a realistic exponential-based model of a query on revoked certificates. This is because, in practice, the number of queries on revoked certificates tend to decrease over time. Our performance analysis is thus more accurate with respect to the CRL size as well as the probability of the query on valid/revoked certificates. Furthermore, we also follow the standardized CSI message sizes in [123] to derive realistic message costs in various revocation schemes.

Our cost analysis framework enhances the one previously proposed in [91, 65]. We summarize the framework of [91, 65] in Section 7.2.3, and describe how we enhance it in Section 7.4.2.

7.2 Preliminaries

Below we provide a brief summary of EV Certificate, CRS/NOVOMODO schemes, and a realistic cost analysis model proposed in [91, 65].

7.2.1 Extended-Validation Certificates (EVC)

EV Certificates (EVCs) are a special type of the X.509 certificate to deal with the problems of phishing [7] and online fraud. It requires a more extensive investigation of the requesting entity by the CA before a certificate is issued. The investigation includes the existence of a domain name and its associated publicly-accessible server which are exclusively owned or controlled by the entity. EVC is important to provide a stronger assurance against some recent attacks on the X.509 certificate usage, such as certification-chaining attack [94], homograph-based phishing [49, 62, 94], and man-in-the-middle based SSLstrip attack [94]. (See Section 2.5.3 for additional background on EV Certificate.)

We remark that, in theory, the CREV schemes proposed here could equally work on the standard X.509 certificate with an included domain name. However, simply accepting a domain name and having the server disseminate the CSI falls short of providing strong assurances on both the certification and the revocation services. The above mentioned attacks in [94, 49, 62], among others, may apply. For our revocation schemes, we need an assurance that a server associated with an EVC principal is a valid server representing the principal as how general public would likely to assume (e.g. `micsOsOft.com` is not `microsoft.com`). Additionally, a principal is bound to perform its best in following the revocation scheme. Our CREV schemes build upon the EVC infrastructure, since it is a widely-accepted standardized premium certificate mechanism that can provide the needed assurance. Any service provided by a PKI that operates under restricting policies at least as secure as those used for EVC can also work with our schemes.

7.2.2 CRS/NOVOMODO

Certificate Revocation Status (CRS) [100] makes revocation more efficient through its periodical release of compact hash chain information. CRS assumes the involvement of CMAE(s), which receive CSI from the CA and handle the verifier's queries.

The basic idea is as follows. Prior to issuing a certificate, the CA chooses a one-way hash function $H()$, and determines a time interval period d and a hash-chain length ℓ. The lifetime of the certificate is: $d\,(\ell{+}1)$. The CA then adds the following to the published certificate: $HashAlgID$ as the identification of $H()$, d, ℓ, issue time, expiration time[5], and two 100-bit values Y and N representing "valid" and "revoked" status, respectively. Y and N are calculated as follows: the CA first generates two secret 100-bit random

[5]The expiration time can actually be derived from: issue time + $d(\ell + 1)$.

numbers Y_0 and N_0, and then computes $Y = H^{\ell}(Y_0)$ and $N = H(N_0)$. The scheme works as follows: on the i-th time interval after the certificate's issuance, where $(1 \leq i \leq \ell)$, the CA submits the following information to CMAE(s): a signed timestamped string containing all serial numbers of issued and not-yet-expired certificates, and 100-bit V_i for each certificate, which indicates whether it has been revoked or not by the current time interval. V_i is set as follows: $V_i = H^{\ell-i}(Y_0)$ if a certificate is still valid, or $V_i = N_0$ if it is revoked. When a CMAE receives a verifier's query, it then sends V_i to the verifier. Using the current time and certificate's issue time, the verifier can check whether a certificate is valid or revoked by comparing $H(V_i)$ with N or comparing $H^i(V_i)$ with Y.[6]

NOVOMODO [101] was later proposed as an improvement on CRS. It suggests the use of SHA-1 as one-way hash function and the avoidance of the CMAE in the centralized NOVOMODO scheme. The hash-chaining mechanism is the same as in CRS. The work [101] also described how to build a distributed NOVOMODO, which makes use of a single vault and an unlimited number of "untrusted responders" to answer validation queries.

The CRS/NOVOMODO schemes have their own limitations. Firstly, it a high in CA-to-CMAE communication cost [114, 5]. Our CREV scheme mitigates this since an EVCP acts as a CMAE *only for its own certificate*. Secondly, CRS/NOVOMODO may require a high verifier's processing cost due to the a large number of repeated hash operations. A hash operation is much more efficient than a public-key operation. However, for a sufficiently fine-grained timeliness guarantee and long certificate lifetime, the total computation cost of all hash operations can be significant particularly for lightweight verifiers. Our CREV scheme reduces the total length of the hash chain to make the verifier more efficient. Lastly, CRS/NOVOMODO increases the CA's storage requirements into $O(n \cdot \ell)$, where n is the number of certificates. This storage requirement can be reduced at the cost of more computation effort by the CA [71], thus trading-off storage with computational cost.

7.2.3 Certificate Revocation Model using Empirical Data

The work [91, 65] aims to determine the CA's strategies in releasing CRL so as to reduce the CA's operational cost. To this end, the work derives a realistic estimate of the number of new revocations between two successive CRL generations as well as the size of CRL on a particular day.

Most previous analysis works on revocation schemes usually assume that a certificate is revoked at the half of its lifetime [177, 84]. In contrast, the work [91, 65] gives a realistic revocation model based on the empirical data of CRLs collected from VeriSign. They found that most of the certificate revocations occur during the early part of the

[6] If a verifier has obtained a valid V_j at an earlier j-th time interval $(j < i)$, the number of hashes required to check the validity is reduced to $i - j$.

certificate's lifetime. In fact, more than 30% of revocations occur within the *first two days* after certificates get issued. In addition, the percentage of revocations decreases as time elapses. This result thus invalidates the assumption about the CRL size taken by most earlier works [177, 84].

The work [91, 65] defines a probability density function (PDF) of certificate revocations over time. Suppose that there are α certificates issued at time X, with uniform issued age β. From time X to $X+\beta$, on average αb of the certificates will be revoked. In [91, 65], the basic time unit between two CRL releases (Δt) is assumed to be of one day. Function $R(t)$ is defined to represent *the revoked percentage*, that is the ratio of the number of revocations that occur in the time interval $[t, t+\Delta t]$, where $X<t<X+\beta$, to the total number of revocations in the time interval $[X, X+\beta]$. The work [91, 65] model the revocation distribution of certificates issued at a particular time with the following exponential PDF[7]:

$$R(t) = ke^{-kt} \tag{7.1}$$

where t is the time, and $k = 0.26$ as the function paramenter which has a good fit to the real observed data. Figure 7.1, reproduced from [91], shows the defined PDF and the actual plot from the empirical data.

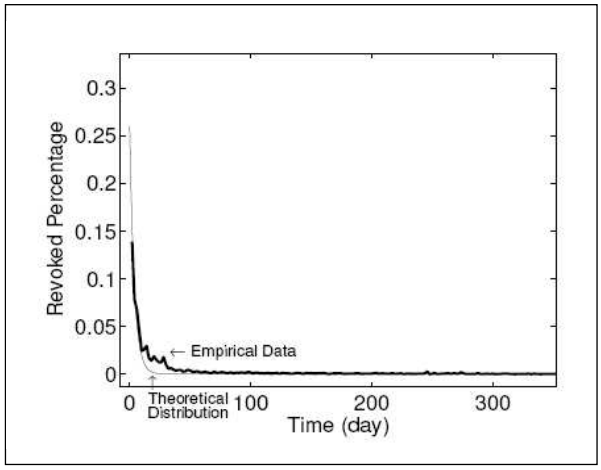


Figure 7.1: The fitted exponential PDF and empirical data for certificate revocations over time (from [91]).

The model in [91, 65] derives a realistic estimate of the CRL data size by deriving the number of revoked entries on a particular day. Three scenarios are considered, namely: a model for certificates issued at different times but with the same issued age β; certificates issued at different times and with different issued ages; and a more general case where the number of certificates generated at a particular time is not constant, but follows a Poisson distribution. The work [91, 65] derives an analytical model for the first and the second scenarios respectively, but use a simulation technique for the third. For our

[7]A Probability Density Function (PDF) is defined by taking $\lim_{\Delta t \to 0}$.

comparison framework in this chapter, we only consider the first scenario. If desired, the framework can be extended to analyze the second scenario in a rather straightforward manner.

The number of *new revocation requests* is derived in [91, 65] as follows. Suppose that v is any time in $(0, \beta]$. Function $f(v)$ is defined as the number of new certificate revocations between day v and day $v+\Delta t$ from all of the valid generations:[8]

$$f(v) = \alpha bR(v) + \alpha bR(v - \Delta t) + \alpha bR(v - 2\Delta t) + \ldots + \alpha bR(v - (n-1)\Delta t) \quad (7.2)$$

where n is the number of certificate generations in time period β, i.e. $n = \lceil \frac{\beta}{\Delta t} \rceil$. As mentioned, [91, 65] assumes $\Delta t = 1$ day. Similarly, the time interval between two certificate generations (ΔX) is also assumed to be 1 day. As a result, v in (7.2) is an integer. Thus, when v is in $(0, \beta]$, we have:

$$f(v) = \alpha bR(1) + \alpha bR(2) + \ldots + \alpha bR(v) = \alpha bke^{-k} \frac{1 - e^{-vk}}{1 - e^{-k}} \quad (7.3)$$

When v is in $(\beta, +\infty)$, which represents the steady-state condition, [91, 65] derives that:

$$f(v) = \alpha bR(1) + \alpha bR(2) + \ldots + \alpha bR(\beta) = \alpha bke^{-k} \frac{1 - e^{-\beta k}}{1 - e^{-k}} \quad (7.4)$$

The *size of CRL* is derived as follows. Let $F(v)$ be the valid cumulative number of revocations from time 1 to v. It is also the size of the CRL on that day. For v in $(0, \beta]$:

$$F(v) = \sum_{t=1}^{v} f(t) = \frac{\alpha bke^{-k}}{1 - e^{-k}} \left[v - \frac{e^{-k}}{1 - e^{-k}} \left(1 - e^{-vk} \right) \right] \quad (7.5)$$

For v in $(\beta, +\infty)$, [91, 65] derives:

$$F(v) = F(\beta) = \sum_{t=1}^{\beta} f(t) = \frac{\alpha bke^{-k}}{1 - e^{-k}} \left[\beta - \frac{e^{-k}}{1 - e^{-k}} \left(1 - e^{-\beta k} \right) \right] \quad (7.6)$$

From (7.5) and (7.6), it can be shown that the size of CRL increases from day 0 until day β. After that, in what we call *the steady-state condition*, the size of CRL becomes constant. Note that some revoked certificates may be expired after they reach their issued lifetimes, and accordingly removed from CRL.

We remark that there are a number of issues with respect to the derivations of $f(v)$ and $F(v)$ in [91, 65], which we discuss and address later in Section 7.4.2.

7.3 CREV Schemes for Lightweight Certificate Revocations

7.3.1 CREV Assumptions and Overview

We make following assumptions between the online transaction servers (EVCPs) and he CA which are reasonable under typical settings of an Internet transaction:

- Synchronized clocks are available to determine the current time on the part of entities involved. CRL, CRS/ NOVOMODO and OCSP also make this assumption.
- There exists an established Service Level Agreement (SLA) between the CA and an EVCP, with a reliable and secure communication channel.

[8] The more recent version of the work [65] uses $N(v)$ instead of $f(v)$ to denote the function.

- The number of verifiers is much more than the number of EVCPs. Moreover, the ratio of number of verifiers to the corresponding EVCPs is high.

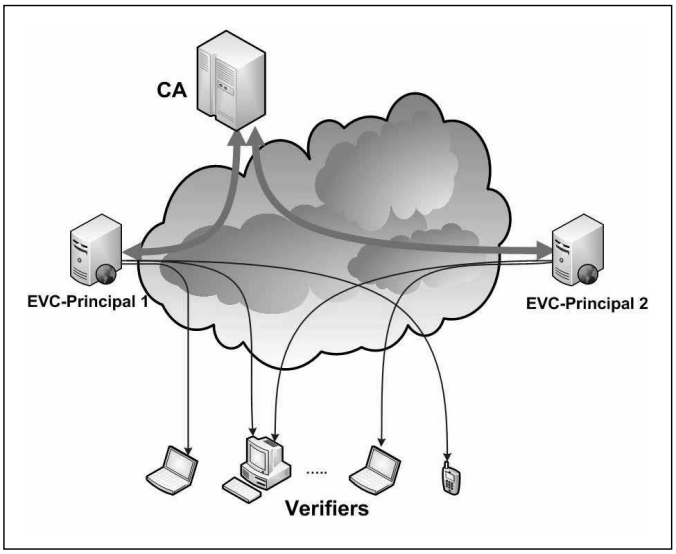


Figure 7.2: CSI communication flow in CREV schemes.

The key to our proposed schemes is the availability of an EVCP's server which acts as a CMAE for its own certificate. Figure 7.2 depicts the CSI communication flow between all parties. Our CREV schemes improve on CRS/NOVOMODO and OCSP by taking advantages of the shown setting. The two revocation schemes are chosen since they issue per-certificate CSI that also can assert the "*liveness*" of a certificate. That is, the CSI can be used as a proof to the verifiers that the certificate is still valid. This is as opposed to the black-listing approach, where non-inclusion in the list is taken as a proof that a certificate is still valid. We need this certificate liveness property since the CSI is obtained by the verifiers from the principal's server, rather than from an authoritative entity (e.g. CA or trusted CMAE).

CREV-I is proposed to improve CRS/NOVOMODO by setting up a *session-based hash-chaining* service between the CA and an EVCP. It aims to take advantage of the hash-chaining technique, but shorten the length of the hash chain for a faster verifier's operation and a reduced CA's storage requirement. We do not deal with amortization techniques such as those based on multidimensional hash chain [71]. Rather, we work at the protocol level to reduce the length of the used hash chain by taking advantage of the applicable EVC setting. The amortization techniques thus can still work together with CREV-I to make its costs even lower. CREV-II is proposed by taking advantage of available online schemes like OCSP. For each (EV) certificate, the CA produces only one OCSP Response per time interval. It also can make use of an optimization technique using group-based CSI [76], which is inspired by the notion of certificate-range in the CRT. Our contribution related to this optimization is thus by showing that such a technique can

fit into the setting, and showing how they can reduce the overheads using the developed cost analysis framework.

To identify the availability of CREV revocation schemes and which type of CREV is in use, we define the following extension to an EV Certificate:

$$\textit{EVC CSI–Access–Point Extension} ::= \textit{Revocation Scheme}, \textit{Transfer Mechanism} \quad (7.7)$$

The fields are as follows:

- **Revocation Scheme**: gives the CREV scheme used.
- **Transfer Mechanism**: defines the CSI transfer mechanism, e.g. HTTP and FTP.

7.3.2 CREV-I: Session-based Hash-Chaining Scheme

The CREV-I scheme extends CRS/NOVOMODO but makes use of the ability of a CA to setup a short-lived session for a periodical *hash-chain token push* to an EVCP. The idea is that each EVCP and the CA establish a secure session using a "3-way Session-Establishment" protocol given below.

Like in OCSP, the CA's availability for a request message may lead to the possibility of Denial of Service (DoS) attacks due to flood of requests to the CA.[9] This attacks can include replaying previously valid request messages to the CA. To deal with this, we require every incoming Hash-Chain Session Establishment Request in CREV-I to be signed by the EVCP. Moreover, we also require the Request to carry T, which is either:

- A timestamp of the current time based on network time protocols.
- A CA's nonce that is accessible by the verifiers from a CA's pre-defined URI. The nonce is regularly updated by the CA at a pre-determined short time interval.

The CA will only process an incoming request with a valid T and the requester's signature. The same requirement on session establishment is also used in the CREV-II scheme.

To support CREV-I, the CA includes the following extension in a published certificate:

- "*EVC CSI-Access-Point Extension*", which is defined in (7.7).
- "*CREV-I Extension*", which contains: URI for CA's nonce (URI_{CA_nonce}), URI for Hash-Chain Session Establishment Reply ($URI_{hashchain\_session}$), and URI for hash-chain token access (URI_{hash_token}).

The scheme works as follows.

CA and EVCP Interaction: Hash-Chain Session Establishment

The following protocol is executed to establish a hash chain session on a currently valid EV Certificate. We define "*Hash-Chain Session Establishment Request*" and "*Hash-Chain Session Establishment Reply*" below to be syntactically similar to OCSP Request

[9] OCSP uses a nonce (original from the requester) to bind a request and a response message. This measure is employed to protect the requester from possible replay attacks. However, there seems to be no mechanism in OCSP to prevent flood-of-requests attack to the responder (see also [113, Section 5]).

and OCSP Response respectively. In the protocol, we assume that $Serial\_No$ and CA's name (CA_ID) are sufficient to identify a unique certificate.[10]

1. $EVCP \rightarrow CA$: "*Hash-Chain Session Establishment Request*"=
   $\langle EVCP\_ID, CA\_ID, Serial\_No, T, nonce_{EVCP}, SigAlgID, Signature\rangle$
   where:
   T = either a timestamp or CA's nonce from URI_{CA_nonce};
   $SigAlgID$ = identification for the signing algorithm.

2. CA: Check T and $Signature$.

3. $CA \rightarrow EVCP$: "*Hash-Chain Session Establishment Reply*"=
   $\langle EstablishmentStatus, CA\_ID, EVCP\_ID, nonce_{EVCP}, Serial\_No,$
   $CertStatus, HashAlgID, d, t_{CREV\_I}, Y, N, SessionStart, SessionExpiry,$
   $SigAlgID, Signature\rangle$
   where:
   $HashAlgID$, d, Y, N = hash chain parameters;
   $SessionStart$ and $SessionExpiry$ = the start and end time of the established session;
   t_{CREV_I}= session's lifetime = $SessionExpiry - SessionStart$.

4. $EVCP$: Check $nonce_{EVCP}$, $Signature$, and that $EstablishmentStatus$ is successful.

5. $EVCP \rightarrow CA$: "*Hash-Chain Session Establishment ACK*"=
   $\langle EVCP\_ID, CA\_ID, Serial\_No, SessionStart, SigAlgID, Signature\rangle$.

6. $EVCP$: Put *Hash-Chain Session Establishment Reply* at $URI_{hashchain\_session}$;
   Establish an association for hash chain updates with CA.

7. CA: Check $SessionStart$ and $Signature$;
   Start providing timely hash-chain token updates until the session expires,
   or the certificate is revoked.

Note that the established session's lifetime t_{CREV_I} is $d\,(\ell_{CREV\_I} + 1)$. Hence, the length of the hash chain in a CREV-I session (ℓ_{CREV_I}) is $\frac{t_{CREV\_I}}{d} - 1$. On the i-th time interval after $SessionStart$, with $(1 \leq i \leq \ell_{CREV\_I})$, the CA then releases the latest hash chain token (V_i) to the EVCP as in CRS/NOVOMODO.

The verifier obtains a Hash-Chain Session Establishment Reply from the specified $URI_{hashchain\_session}$, and V_i from URI_{hash_token}. It then validates the CA's $Signature$ in

[10] OCSP [113] makes use of a serial number together with $HashAlgID$, $H(CA\_ID)$ and *H(CA's public-key)* due to the presumed lack of strong uniqueness guarantee for CA names.

Hash-Chain Session Establishment Reply, and then determines whether the EVC is still valid based on the values of V_i, Y, and N.

The length of hash chain used in CREV-I is ℓ_{CREV_I} which is much smaller than that in CRS/NOVOMODO. As a result, the workload for the verifier's repeated hash operations is much reduced. By the same token, the CA now stores and keep tracks of a much smaller (session-wide) hash chain. This is in contrast with the hash chain used in CRS/NOVOMODO, which is constructed for the whole certificate's lifetime.

A formal analysis of the session establishment protocol using MPKI-BAN Logic (the subject of Chapter 8) is given in the Appendix E.

7.3.3 CREV-II: Session-based Online Status Scheme

The CREV-II scheme takes advantage of a CA's ability to support an online status notification service such as OCSP. CREV-II improves on CREV-I by providing a *CA's direct (signed) proof* of certificate goodness to EVCPs.[11] Given OCSP as a standardized scheme for online status notification, we here simply assume the online status assurance in the form of OCSP Response. Thus, CREV-II allows the CA and the verifiers to take advantage of their existing OCSP infrastructures. Unlike the original OCSP, the CA in CREV-II provides the service only to EVCPs, which in turn make the CSI available to the respective verifiers. This new setting also allows for a workable economic model in which the CA charges EVCPs, rather than the verifiers, for the CSI services.

To support CREV-II, the CA includes the following extensions in a certificate that it publishes:

- "*EVC CSI-Access-Point Extension*" defined in (7.7).
- "*CREV-II Extension*" containing: URI for CA's nonce (URI_{CA_nonce}) and URI for the latest OCSP Response (URI_{latest_OCSP}).

The CREV-II scheme works as follow.

The CA and EVCP Interaction: Status Subscription Establishment

To establish a CSI subscription session, the CA and EVCP perform the following protocol:

1. $EVCP \rightarrow CA$: "*Status Subscription Establishment Request*"$= \langle EVCP\_ID,$ $CA\_ID, Serial\_No, T, nonce_{EVCP}, d_{EVCP}, t_{EVCP}, SigAlgID, Signature\rangle$

   where:

   T = either a timestamp or CA's nonce as in CREV-I;

   d_{EVCP} = proposed time interval between two OCSP Response messages;

   t_{EVCP} = proposed lifetime of the established session.

[11] In reporting a certificate's status, an online service is assumed in this thesis to provide its timeliest possible revocation guarantee as permitted by its policy.

2. CA: Check T and $Signature$.

3. $CA \rightarrow EVCP$: "*Status Subscription Establishment Reply*"$=$
   $\langle EstablishmentStatus, CA\_ID, EVCP\_ID, Serial\_No, nonce_{EVCP}, CertStatus,$
   $d_{CA}, \ell_{CA}, SessionStart, SessionExpiry, SigAlgID, Signature\rangle$.
   The session's lifetime (t_{CREV_II}) is $= SessionExpiry - SessionStart = d_{CA} \cdot \ell_{CA}$.

4. $EVCP$: Check $nonce_{EVCP}$, $Signature$, and that $EstablishmentStatus$ is successful.

5. $EVCP \rightarrow CA$: "*Status Subscription Establishment ACK*"$=$
   $\langle EVCP\_ID, CA\_ID, Serial\_No, SessionStart, SigAlgID, Signature\rangle$.

6. $EVCP$: Establish an association for status updates with CA.

7. CA: Check $SessionStart$ and $Signature$;
   Start providing timely online status updates to EVCP.

Upon completion of Step 7, the CA starts delivering OCSP Response messages to the EVCP at every d_{CA} time interval for ℓ_{CA} times; or until the certificate is revoked. Upon receipt of the latest OCSP Response, the EVCP puts it on URI_{latest_OCSP}.

Note that the periodically released OCSP Responses contain no requester's challenge (nonce). As such, the EVCP and the verifier must verify the freshness of a Response message by checking the values of `thisUpdate` and `nextUpdate` to be current. The CA must properly set a Response's validity period (`nextUpdate` - `thisUpdate`) to d_{CA}. Note that the CA can pre-produce the periodically released OCSP Responses.

If d_{CA} for multiple EVCPs is short, e.g. 1 minute, a further optimization on OCSP Response is possible. The CA can issue an OCSP Response that carries the statuses of *multiple certificates*. The validity periods of all the certificate statuses are set uniformly.

If the EVC serial number is sequential and that all EVCs make use of CREV-II, we can even apply a *range optimization* technique which is based on the notion of certificate range in CRT [75]. The technique is similar to the one proposed by Koga et al. [76]. Our use of the technique here show how it can be utilized in the proposed certification setting. We also show later how well the optimization can reduce the overheads. The technique works as follows. The CA keeps track of valid certificates as a set of *good certificate ranges*. Each good certificate range is defined as $[SN_a, SN_b]$ indicating that all certificates with serial number SN_i, where $SN_a \leq SN_i \leq SN_b$, are all valid. For an optimal range construction (resulting in the smallest number of ranges), we also require

that $SN_a - 1$ and $SN_b + 1$ are revoked (or possibly expired). If a certificate is revoked, the CA sends a single OCSP Response notifying the revoked status of that certificate. For a valid certificate with serial number SN_v, the CA sends an OCSP Response with a good certificate range where $SN_a \leq SN_v \leq SN_b$. Given that the majority of certificates are good ($\sim 90\%$ in our evaluation), this optimization can significantly reduce the CA's signature operations as shown later in Section 7.4.4.

7.4 Analysis, Evaluation and Comparison of CREV Schemes

7.4.1 Security Analysis of CREV Schemes

The security of CREV schemes relies mainly on the underlying revocation schemes, namely CRS/NOVOMODO and OCSP. Provided that all the entities run with relatively synchronized clocks, the freshness of the released hash tokens (in CREV-I) and OCSP Responses (in CREV-II) can be established. The *Hash-Chain Session Establishment Reply* (in CREV-II) is signed by the CA. It behaves as a kind of "mini certificate" which carries information about the CA's hash chain parameters. The CREV-II scheme takes advantage of the OCSP Response model, where the verifier only needs to check the validity of the CA's signature and timestamp.

We project that the CREV schemes are practical with near a real-time timeliness guarantee from ten to one minute. A ten minute guarantee may already be considered reasonably short-lived in many environments. A one minute guarantee could be even considered as being "indistinguishable" from a real-time service due to potential clock time differences among the involved entities [101]. The recency requirements in our schemes are thus set by the CA and EVCP, and not the verifier (as acceptor). As such, the schemes still fail to address an issue pointed out in [128]. Yet, with a near real-time timeliness guarantee, a verifier can proceed with the transactions only if it feels satisfied with the offered more fine-grained timeliness guarantee.

7.4.2 A Framework for Performance Analysis

Our performance analysis framework extends the model given in [91, 65]. A novelty of our framework is that it includes more real-world features and cost estimates.

Certification System Assumptions

Our analysis is primarily concerned with measuring the overheads of revocation schemes in a *steady-state condition* with a single CA certification system. Since our objective is to provide a near real-time freshness guarantee, we make use of *minute* as our unit of time. We make the following steady-state assumptions:

- The total number of principals and the corresponding valid certificates (N) is constant.
- Among the N certificates, N_{EV_cert} are EV certificates and N_{Reg_cert} are Regular (non-EV) certificates. We have $N_{EV\_cert} + N_{Reg\_cert} = N$.
- Certificates have the same lifetime, which is β days $= \beta \times 1,440$ minutes.[12]
- The time interval between two successive CRL releases (Δt) is constant. The unit of Δt is minute(s).
- Certificate issuance takes place at a constant rate. The time interval between two successive certificate generations (ΔX) is the same as Δt.
- There are $b \cdot N$ certificates revoked in the span of β days. Revocation affects both N_{EV_cert} and N_{Reg_cert} proportionally.
- A new valid certificate is issued immediately to replace a revoked one. This can take place in practice by assuming that a principal always submits a revocation notice together with a replacement request to prevent any operational interruption.

Since we consider a certification system with uniform certificate lifetime β, the steady-state condition takes place after β days from the *first* certificate generation in the system, i.e. the time interval $(\beta, +\infty)$.

By considering a system where $\Delta X = \Delta t = 1$ minute as an example of our evaluation scenario, we thus cover the situation where the certificate issuance and the CRL release take place continuously.[13] As mentioned earlier, most earlier works [177, 84] make an assumption that a certificate is assumed to be revoked at half of its lifetime ($\frac{\beta}{2}$). Such an assumption is however crucial to the performance analysis of revocation schemes as it affects, among others, the estimate of the CRL size. Here, we define our improved model which is more realistic.

An Improved Model and Analysis

The work [91] (summarized in Section 7.2.3) is the first to suggest a realistic revocation model, which makes use of real empirical data of CRLs. Our analysis employs the same approach as [91, 65] but increases the realistic features as in the following aspects.

1. Revision of the CRL Size Estimate

The work [91, 65] always assume that $\Delta X = \Delta t = 1$ day. As such, there are $\alpha = \frac{N}{\beta}$ certificates issued on day X, which are valid between day X and day $X+\beta$. For certificates

[12] The work [91, 65] also considers a scenario with certificates having several ages.

[13] Given our derivation of $h_g(v)$ (Eq. 7.15) and $Rev\_Entries_g$ (Eq. 7.16), it is fairly straightforward to generalize the CRL size estimate when $\Delta X \neq \Delta t$.

issued at time X, $R(t)$ is defined to represent the revoked percentage. It is modeled as an exponential PDF of $R(t) = ke^{-kt}$, where t is the time and the parameter $k = 0.26$.

The work [91, 65] calculates the CRL size estimate using the derivation shown in Eqs. 7.2–7.6. We note two issues with this derivation. Firstly, their calculations make use of $f(v)$ in [91] (or $N(v)$ in [65]) to obtain the number of new certificate revocations between day v and day $v+\Delta t$. The functions however use a *discrete summation* on the PDF which is only an approximation (see Eqs. 7.2–7.4). In our work, we derive the number of new certificate revocations using *integration*, which is a more proper method, on the same PDF. This difference becomes important when ΔX and Δt is much less than one day. Simply applying the summation technique as in [91, 65] would derive the number of revoked entries that is *much larger* than that of the total revoked certificates $N \cdot b$ (see Section 7.4.4 for a sample figure based on two selected evaluation scenarios). Secondly, in the steady-state calculation (Eq. 7.4), $f(v)$ or $N(v)$ actually measures the the number of revocations for days $[1, \beta + 1]$ instead of $(0, \beta)$.

We reformulate $F(v)$ where $\Delta X = \Delta t = 1$ day by first defining $g(v)$ to replace $f(v)$ in [91] or $N(v)$ in [65]. Function $g(v)$ is defined as the number of new revocations between day $v-\Delta t$ and day v from all valid generations. When v is in $[0, \beta]$, we have:

$$g(v) = \sum_{i=1}^{v} \alpha b \int_{i-1}^{i} R(t)\, dt = \alpha b \int_{0}^{v} ke^{-kt}\, dt = \alpha b \left[-e^{-kt} \right]_0^v = \alpha b \left(1 - e^{-kv}\right). \quad (7.8)$$

For the steady-state condition, i.e. when $v \in (\beta, +\infty)$, we have:

$$g(\beta) = \alpha b \int_{0}^{\beta} ke^{-kt}\, dt = \alpha b \left[-e^{-kt} \right]_0^\beta = \alpha b \left(1 - e^{-k\beta}\right). \quad (7.9)$$

$F(v)$, representing the cumulative number of certificate revocations from day 0 to day v, with $v \in [1, \beta - 1]$ becomes:

$$F(v) = \sum_{i=1}^{v} g(i) = \sum_{i=1}^{v} \alpha b \int_{0}^{i} ke^{-kt}\, dt = \alpha b \left[v - e^{-k}\, \frac{1 - e^{-kv}}{1 - e^{-k}} \right]. \quad (7.10)$$

For $v \in [\beta, +\infty)$, which includes the steady-state condition, we thus have:

$$Rev\_Entries = F(\beta - 1) = \sum_{i=1}^{\beta-1} g(i) = \alpha b \left[(\beta - 1) - e^{-k}\, \frac{1 - e^{-k(\beta-1)}}{1 - e^{-k}} \right]. \quad (7.11)$$

Note that in Eq. 7.11, we make use of $F(\beta - 1)$ instead of $F(\beta)$ in Eq. 7.6 as in [91, 65]. This is since there are only $\beta - 1$ valid certificate generations, and not β. The generation on day v has no revoked certificates yet, and the generation on day $v-\beta$ is expired on that day, hence leaving only $\beta - 1$ generations from day $v-1$ to $v-\beta+1$.

We also find out that there is a simpler and more elegant way of calculating the CRL

size on day v, which is described as follows. First, we define $h(i)$ as the total number of (non-expired) revoked certificates from certificates generated in i generation(s) *prior to* the current generation:

$$h(i) = \alpha b \int_0^i R(t)\, dt. \tag{7.12}$$

$h(i)$ thus represents the cumulative distribution function of $R(t)$. The number of entries in CRL during the steady-state condition (with $\beta-1$ certificate generations counted) is:

$$Rev\_Entries = \sum_{i=1}^{\beta-1} h(i) = \sum_{i=1}^{\beta-1} \alpha b \int_0^i R(t)\, dt = \alpha b \left[(\beta - 1) - e^{-k} \frac{1 - e^{-k(\beta-1)}}{1 - e^{-k}} \right], \tag{7.13}$$

yielding the same result as Eq. 7.11. In our generalization below, we will make use of $h(i)$.

2. Generalization of the CRL Size with Certificate Generation (ΔX) and CRL Issuance (Δt) in Minutes

In our performance evaluation, we set the revocation timeliness guarantee of 1 and 10 minutes. To deal with this, we need to generalize the calculation of the number of CRL entries where $\Delta X = \Delta t = \delta$ minutes.

Let M denote the number of minutes in a day, that is M=1,440. We then define a constant $\lambda = \frac{\delta}{M}$. The number of certificate issued per generation (α_g) is now:

$$\alpha_g = \frac{N\lambda}{\beta} = \alpha\lambda. \tag{7.14}$$

We generalize $h(v)$ into $h_g(v)$ as follows:

$$h_g(v) = \alpha_g b \int_0^i R(t)\, dt = \alpha_g b \int_0^i k e^{-kt} = \frac{N\lambda}{\beta}\, b \left(1 - e^{-kv}\right). \tag{7.15}$$

The number of revoked entries during the steady-state condition where $v \in (\beta, +\infty)$ with $\Delta X = \Delta t = \frac{\delta}{M}$ day is:

$$\begin{aligned} Rev\_Entries_g &= \sum_{i=1}^{\eta} h_g(\lambda i) = \sum_{i=1}^{\eta} \alpha_g b \int_0^{\lambda i} R(t)\, dt \\ &= \frac{N\lambda}{\beta}\, b \left[\left(\frac{\beta}{\lambda} - 1 \right) - e^{-k\lambda} \frac{1 - e^{-k(\beta-\lambda)}}{1 - e^{-k\lambda}} \right]. \end{aligned} \tag{7.16}$$

where $\eta = \frac{\beta M}{\delta} - 1$ denotes the number of counted generations.

3. Different Query Rates on EV and Regular Certificates

In our evaluation, we distinguish the query rate between EV and Regular certificates. This is because we expect EV certificates to be employed mostly by servers providing Internet transactions. Thus, we assume that the average daily query on an EV certificate

($q_{per\_EV\_daily}$) far exceeds that for a Regular one ($q_{per\_Reg\_daily}$) by a ratio μ. That is:

$$\frac{q_{per\_EV\_daily}}{q_{per\_Reg\_daily}} = \mu. \tag{7.17}$$

The total number of daily query on all EV certificates is: $Q_{EV\_daily} = N_{EV\_cert} \times q_{per\_EV\_daily}$. Likewise, the total daily query on Regular certificates is: $Q_{Reg\_daily} = N_{Reg\_cert} \times q_{per\_Reg\_daily}$. Hence, we have the total daily query:

$$Q_{daily} = Q_{EV\_daily} + Q_{Reg\_daily}. \tag{7.18}$$

4. More Realistic Models for Query on Revoked Certificates

Determining whether a certificate being queried is either revoked or valid is a key factor in calculating the computational costs of a hash-chaining based scheme like CRS, NOVOMODO or CREV-I. This is since there is a great difference in cost between verifying a valid certificate and a revoked one. To derive realistic probabilities, we first need to have realistic model(s) for query on the revoked (but not-yet-expired) certificates.

The function $S(t)$ is defined as the probability that a certificate will still be queried in the time interval $[(t-1)\delta,\, t\delta\,]$ minutes after it is first known to be revoked.[14] Here, we take into account the fact that certificates are used for different purposes, such as for online transactions and code/document signing. As such, one should also have different query models to suit these different purposes. We define the following two query models:[15]

Model A (Online-transaction certificates): This applies to certificates used for online transactions where a revoked certificate will quickly cease to be referred to. We define the following probability function for this model (shown in Figure 7.3):

$$S_A(t) = e^{-k_A\,\lambda\,t} \tag{7.19}$$

with the parameter k_A set to 0.95.

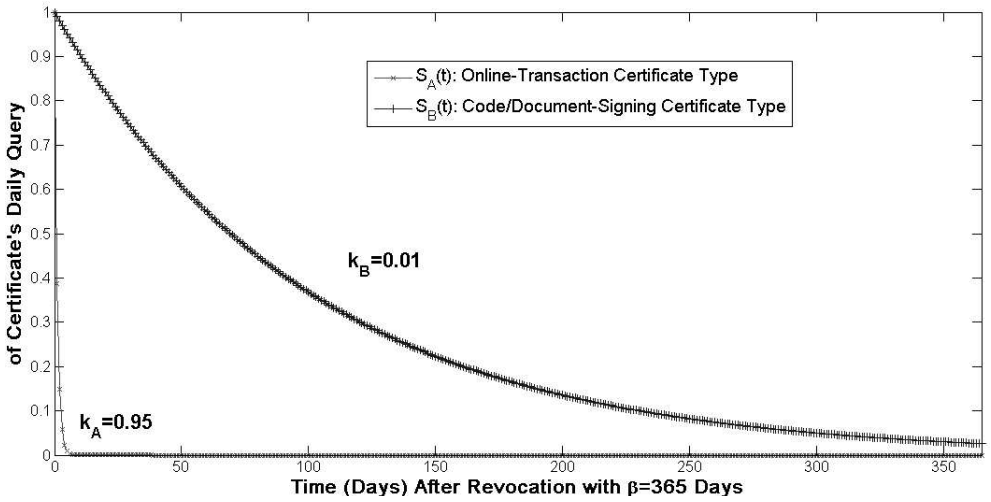
Model B (Code/document-signing certificates): This applies to certificates used to provide non-online services, such as the signing of documents/software that may be distributed in an off-line manner (e.g. using a physical media). In this type of certificates, a revoked certificate will still be referred to for a longer time period following its revocation. We model its probability function as:

$$S_B(t) = e^{-k_B\,\lambda\,t} \tag{7.20}$$

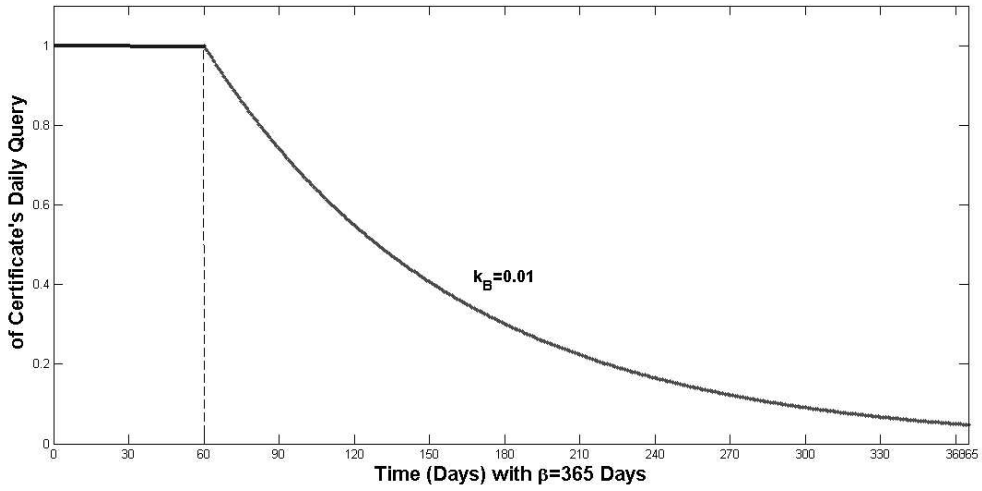
with $k_B = 0.01$ (see Figure 7.3).

[14] Unlike $R(t)$, $S(t)$ indicates the *probability value* at time t rather than a PDF. As such, in Eq. 7.21 we make use of summation over $S(t)$ as opposed to integration in deriving the total number of queries on the revoked certificates.

[15] A CA can have a different number of certificates classes. However, we assume that these classes can be mapped into the two defined query models.



(a) Two probability functions $S_A(t)$ and $S_B(t)$ are shown in the time interval $[0, \beta$-days) after the certificate's revocation. The function $S_A(t)$ is shown as an L-shaped curve close to the axes.



(b) A sample probability function (of $S_B(t)$ model) on a certificate that is revoked on day 60. Prior to the certificate's revocation, its query probability is 1.0.

Figure 7.3: Probability functions for query on a revoked certificate.

5. Derivation of the Probabilities of Query on Revoked and Valid Certificates

Our objective here is to derive the probabilities of whether a query from a verifier is issued on a valid certificate (Pr_{valid}) and a revoked one (Pr_{rev}). We derive these two probabilities as follows.

First, recall that the daily average query received by CA/CMAE is Q_{daily} (see Eq. 7.18).[16] Now, we derive $Q_{rev\_\beta\_days}$ which is the total number of queries on *all revoked* certificates (until their expirations) *throughout* β days under the steady-state condition. Since we have two certificate types (i.e. EV and Regular) and two query models on the revoked certificates (i.e. model A and B), we define $Q_{rev\_\beta\_days}$ for $x \in \{EV, Reg\}$ and $y \in \{A, B\}$ as follows:

$$Q_{rev\_\beta\_days\_\langle x\rangle\_\langle y\rangle} = \sum_{i=1}^{\eta} \Big(N_{\langle x\rangle\langle y\rangle} \cdot b \cdot \int_{\lambda(i-1)}^{\lambda i} R(t)\, dt \Big) \cdot \Big(\sum_{j=1}^{\frac{\beta}{\lambda}-i} S_{\langle y\rangle}(j) \cdot q_{per\_\langle x\rangle\_per\_interval} \Big) \quad (7.21)$$

where $\eta = \frac{\beta M}{\delta} - 1$ denotes the number of counted certificate generations, $\lambda = \frac{\delta}{M}$, and $q_{per\_\langle x\rangle\_per\_interval} = \lambda \cdot q_{per\_\langle x\rangle\_daily}$ denotes the number of queries on a certificate of type x in the time interval of δ minutes. The placeholders $\langle x\rangle$ and $\langle y\rangle$ in Eq. 7.21 are to be replaced with the relevant options in each combination case.

Summing up all possible combinations, we have:

$$Q_{rev\_\beta\_days} = \sum_{x \in \{EV, Reg\}} \sum_{y \in \{A,B\}} Q_{rev\_\beta\_days\_\langle x\rangle\_\langle y\rangle} \quad (7.22)$$

The total number of query on *all* non-expired certificates, both valid and revoked ones, for β days is $Q_{all\_\beta\_days} = Q_{daily} \cdot \beta$. Hence, we have the total number of queries on *all valid* certificates throughout β days under the steady-state condition: $Q_{valid\_\beta\_days} = Q_{all\_\beta\_days} - Q_{rev\_\beta\_days}$. Finally, the probabilities of whether a status query is issued on a valid certificate (Pr_{valid}) and a revoked certificate (Pr_{rev}) are:

$$Pr_{rev} = \frac{Q_{rev\_\beta\_days}}{Q_{all\_\beta\_days}} \qquad \text{and} \qquad Pr_{valid} = \frac{Q_{valid\_\beta\_days}}{Q_{all\_\beta\_days}}. \quad (7.23)$$

6. Taking Realistic CSI Message Sizes into Account

Many previous analysis works [177, 84] take a simplified (or minimalistic) approach with respect to the CSI message sizes of standardized revocation schemes. In practice, however, the CSI also include various auxiliary information. To achieve a realistic analysis, we follow the CSI message sizes in [123], which utilize a BER viewer for all ASN.1 standardized data structures. Furthermore, we also always assume SHA-1 for the hash algorithm and SHA-1 with RSA for the signature algorithm in all the examined schemes.

[16] Q_{daily} denotes the number of daily CSI look-ups received by the CA/CMAE. Given the short timeliness guarantees in our setting, we assume that all queries are always consulted with the CA/CMAE. Thus, no CSI cache is employed by a verifier. This is reasonable since we set the verifier's daily query (Q_{Ver_daily}) to 30, which is much less than the number of daily CSI updates (1,440 and 144). As a result, we can employ Q_{daily} for calculating query overheads in both the batch-updating and the online schemes.

Performance Comparison Metrics

We use the following notations to denote various cost factors for a particular scheme: Ovh_A= computation time needed by entity A (in seconds), Bw_{A-B} = network bandwidth needed from entity A to B (in MB), and $Stor_A$ = storage needed on A (in MB).

In comparing the schemes, we use the following metrics:

1. Certificate creation cost: Ovh_{CA}.
2. Update costs: Ovh_{CA}, Ovh_{CMAE} (Ovh_{EVCP}) and $Bw_{CA-CMAE}$ ($Bw_{CA-EVCP}$).
3. Query costs: Ovh_{CA}, Ovh_{CMAE} (Ovh_{EVCP}), $Bw_{CMAE-Ver}$ ($Bw_{EVCP-Ver}$) and Ovh_{Ver}.
4. Storage requirement at one point in time (during the steady-state condition): $Stor_{CA}$ and $Stor_{CMAE}$ ($Stor_{EVCP}$).
5. Timeliness: the revocation timeliness guarantee, which also represents the *window of vulnerability* of the revocation scheme.

For metrics (1)–(3), we measure the total *daily* overheads. A summary of the notations used for the performance comparison is given in Table 7.1.

7.4.3 Performance Comparison

Here, we measure the overheads of CRL scheme (with a CMAE used), OCSP (with the CA as a Responder), CRS (with a CMAE used), and our two CREV schemes. For bandwidth calculation, we do not consider the cost due to the underlying network transfer mechanism(s), i.e. HTTP, TCP, or IP. For CRL and CRS which makes use of CMAE, we assume of one CMAE involved respectively. We use L_M to denote the length of message portion M.

CRL (with a CMAE)

We have derived the number of revoked (but non-expired) certificates where $\Delta X = \Delta t = \delta$ minutes as a constant number $Rev\_Entries_g$ (see Eq. 7.16). CRL size is thus defined as: $L_{CRL} = L_{CRL\_fields} + \lfloor Rev\_Entries_g \rfloor \cdot L_{CRL\_entry}$, with L_{CRL_fields}=400 bytes as the message length of CRL header and signature, and L_{CRL_entry}=39 bytes as the length of each entry in the CRL data [123]. With $U = \frac{M}{\delta}$ as the total number of CRL updates in a day, the daily update costs are: $D\_Bw_{CA-CMAE} = U \cdot L_{CRL}$, $D\_Ovh_{CA} = U \cdot C_{sign}$, and $D\_Ovh_{CMAE} = U \cdot C_{verify}$.

The daily query costs of the CRL scheme are: $D\_Bw_{CMAE-Ver}$ (for CMAE)= $Q_{daily} \cdot L_{CRL}$, $D\_Bw_{CMAE-Ver}$ (for Ver)= $Q_{Ver\_daily} \cdot L_{CRL}$, $D\_Ovh_{CA} = 0$, $D\_Ovh_{CMAE} = 0$, and $D\_Ovh_{Ver} = Q_{Ver\_daily} \cdot C_{verify}$, with Q_{Ver_daily} as the verifier's average number of daily queries to the CMAE/CA.

Symbol	Description	Unit	Value(s)
Parameters for Certification & Revocation System:			
N	No. of valid (non-revoked) and not-yet-expired certs	-	250,000
N_{EV_cert}	Number of EV certificates out of N	-	50,000
N_{Reg_cert}	Number of Regular (non-EV) certificates out of N	-	200,000
β	Lifetime of certificate	days	365
$\Delta X = \delta$	Time interval between two successive cert generations	mins	1 | 10
$\Delta t = \delta$	Time interval between two successive CRL releases	mins	1 | 10
b	Percentage of certificates revoked	-	0.1 = 10%
M	Number of minutes in a day	-	1,440
λ	Factor of δ/M	-	δ/M
α_g	No. of certificates issued per certificate generation	-	$N\lambda/\beta$
η	No. of cert generations counted to calculate CRL size	-	$\frac{\beta}{\lambda} - 1$
$d = \delta$	Time interval for periodic hash-token release in CRS/NOVOMODO/CREV-I	mins	1 | 10
ℓ	Hash chain length which determines certificate lifetime in CRS/NOVOMODO/CREV-I	-	$\frac{\beta}{\lambda} - 1$
U	Total number of CRL and hash-token releases per day	-	M/δ
t_{CREV_I}	Lifetime of hash chain in a CREV-I session	hrs	3 | 12
t_{CREV_II}	Lifetime of a status subscription session in CREV-II	hrs	3 | 12
k	Parameter for realistic certificate revocation PDF	-	0.26 (Ref. [91, 65])
k_A	Parameter for class-A query model on revoked certs	-	0.95
k_B	Parameter for class-B query model on revoked certs	-	0.01
N_{EV_A}	No. of EV certificates with class-A query model	-	40,000
N_{EV_B}	No. of EV certificates with class-B query model	-	10,000
N_{Reg_A}	No. of Regular certificates with class-A query model	-	120,000
N_{Reg_B}	No. of Regular certificates with class-B query model	-	80,000
μ	Ratio between daily query on an EV and Regular certs	-	10,000:1
$q_{per\_EV\_daily}$	Average daily queries on an EV certificate	-	1,000,000
$q_{per\_Reg\_daily}$	Average daily queries on a Regular certificate	-	100
$q_{per\_cert\_daily}$	Average daily queries on a certificate	-	due to Eq. 7.18
Q_{Ver_daily}	Average daily queries to CA/CMAE from a verifier	-	30
Basic Cryptographic Operation's Overheads and CSI Size:			
C_{sign}	Cost of a digital signature generation (RSA-1024)	ms	1.48
C_{verify}	Cost of a digital signature verification (RSA-1024)	ms	0.07
C_{hash}	Cost of computing a hash (SHA-1)	μs	0.40
L_{hash}	Length of hash output (SHA-1)	bytes	20
Revocation Scheme's Cost Factors:			
Bw_{A-B}	Bandwidth requirement from entity A to B	MB	-
Ovh_A	Computation overheads on entity A	sec	-
$Stor_A$	Storage requirement on A for CSI dissemination	MB	-
$D\_\langle Cost \rangle$	Total daily (bandwidth, computation, storage) costs	MB|sec	-

Table 7.1: Notations used in the performance analysis. Value(s) column shows the selected values for the two investigated scenarios.

The storage requirements are: $Stor_{CA} = L_{CRL}$, $Stor_{CMAE} = L_{CRL}$. The daily cost of certificate creation is: $D\_Ovh_{CA} = \frac{N}{\beta} \cdot C_{sign}$. Finally, the timeliness guarantee is δ minutes.

OCSP (with CA as Responder)

Below, we give the overheads of OCSP where the CA functions as the OCSP Responder. Here we assume a setting where a nonce is used to bind an OCSP Response with the corresponding Request. There is no update cost between the CA and CMAE, since no CMAE is involved.

With $L_{OCSP\_Response}$=459 bytes as the length of OCSP Response [123], the daily query costs (due to status reply) are: D_Bw_{CA-Ver} (for CA)= $Q_{daily} \cdot L_{OCSP\_Response}$, D_Bw_{CA-Ver} (for Ver)= $Q_{Ver\_daily} \cdot L_{OCSP\_Response}$, $D\_Ovh_{CA} = Q_{daily} \cdot C_{sign}$, and $D\_Ovh_{Ver} = Q_{Ver\_daily} \cdot C_{verify}$.

The storage requirement is $Stor_{CA} = 0$, since the CA needs no additional data structure (apart from the original certificate records). The daily cost of certificate creation is $D\_Ovh_{CA} = \frac{N}{\beta} \cdot C_{sign}$. The timeliness guarantee of OCSP scheme can be close to zero when desired.

CRS (with a CMAE)

CRS scheme is proposed to work with a CMAE [100]. Here, we set $d = \delta$ minute(s). Since the certificate lifetime is βM minutes, we have the length of the hash chain $\ell = \frac{\beta M}{\delta} - 1$. Here, we assume that the CA stores the whole hash chain of all the valid certificates in its storage. An amortization technique such as [71] would reduce its storage requirements, but with the CA's additional online processing.

We use L_{CRS_fields}=161 bytes to denote the length of the CA's timestamp and signature, and L_{Serial_No}=7 bytes to denote the length of a certificate's serial number [123]. The bandwidth cost for update between CA and CMAE is: $Bw_{CA-CMAE} = L_{CRS\_fields} + (N + \lfloor Rev\_Entries_g \rfloor) \cdot (L_{Serial\_No} + L_{hash})$. With $U = \frac{M}{\delta}$, the total daily update costs are: $D\_Bw_{CA-CMAE} = U \cdot Bw_{CA-CMAE}$, $D\_Ovh_{CA} = U \cdot C_{sign}$, and $D\_Ovh_{CMAE} = U \cdot C_{verify}$.

For query, the overheads incurred on the verifier depends on the following conditions. If the certificate is revoked, then $Ovh_{Verifier\_rev} = C_{hash}$. If the certificate is valid, the overhead varies according to the number of hash operations required. The best case occurs when $i = 1$ or when the verifier previously has checked V_{i-1}, thus $Ovh_{Verifier\_valid\_best} = C_{hash}$. The worst case happens when $i = \ell$, in which the verifier needs to perform ℓ hash operations, resulting in $Ovh_{Verifier\_valid\_worst} = \ell \cdot C_{hash}$. The average number of hash computations is $\frac{\ell+1}{2}$, hence $Ovh_{Verifier\_valid\_avg} = \frac{\ell+1}{2} \cdot C_{hash}$. The verifier's average query cost which encompasses both the revoked and the valid cases

is therefore $Ovh_{Verifier\_Avg} = Pr_{rev} \cdot Ovh_{Verifier\_rev} + Pr_{valid} \cdot Ovh_{Verifier\_valid\_avg}$, with Pr_{rev} and Pr_{valid} defined in Eq. 7.23.

The corresponding daily query costs are: $D\_Bw_{CMAE-Ver}$ (for CMAE)= $Q_{daily} \cdot L_{hash}$, $D\_Bw_{CMAE-Ver}$ (for Ver)= $Q_{ver\_daily} \cdot L_{hash}$, $D\_Ovh_{CA} = 0$, $D\_Ovh_{CMAE} = 0$, and $D\_Ovh_{Verifier\_Avg} = Q_{ver\_daily} \cdot Ovh_{Verifier\_Avg}$.

For storage calculation, note that the CA can remove the subchains it has released. Hence, for a certificate issued i generations prior to the current time, the CA only needs to store $(\ell - i)$ hash tokens together with N_0. Thus, the storage requirements in CRS are: $Stor_{CA} = \sum_{i=1}^{\ell} \frac{N\lambda}{\beta} \cdot (i+1) \cdot L_{hash} + \lfloor Rev\_Entries_g \rfloor \cdot L_{hash} = \frac{N\lambda}{\beta} \cdot \frac{\ell^2+3\ell}{2} \cdot L_{hash} + \lfloor Rev\_Entries_g \rfloor \cdot L_{hash}$, and $Stor_{CMAE} = Bw_{CA-CMAE}$. For each certificate, the total cost for generating hash chains (from Y_0 and N_0) is: $(\ell + 1) \cdot C_{hash}$. However, this hash generation parts can be pre-computed. The online overheads which we are concerned here is $Ovh_{CA} = C_{Sign}$, and the corresponding daily cost is $D\_Ovh_{CA} = \frac{N}{\beta} \cdot C_{Sign}$. The timeliness guarantee of CRS is $d = \delta$ minutes.

CREV-I

For the cost calculation of the two CREV schemes, we assume that all certificates are of EV certificate type. Thus, $N = N_{EV\_cert}$ and $q_{per\_cert\_daily} = q_{per\_EV\_daily}$. Later in the evaluation scenarios, we also consider an additional setting where CSI disseminations of all Regular certificates are handled by a designated server acting as the (proxy) CMAE.

CREV-I establishes a short-term session of t_{CREV_I} minutes between the CA and an EVCP. To achieve the timeliness guarantee of δ minutes, we set the hash-chain update interval in CREV-I (d) to δ minutes. Assuming an uninterrupted service, each EVCP thus performs $S = \frac{M}{t_{CREV\_I}}$ session establishments daily. Hence, each EVCP receives $U = S \cdot \ell_{CREV\_I} = \frac{M}{\delta} - S$ hash-token updates in a day. We denote the length of Hash-Chain Session Establishment Reply as $L_{CREV\_I\_Reply} = 605$ bytes, and the length of all messages in a session establishment as $L_{CREV\_I\_Msgs} = 1{,}615$ bytes.

The total daily update costs (comprising both the session establishments and the hash-token updates) are: $D\_Bw_{CA-EVCP}$ (for CA) $= N \cdot (S \cdot L_{CREV\_I\_Msgs} + U \cdot L_{hash})$, $D\_Bw_{CA-EVCP}$ (for EVCP)= $S \cdot L_{CREV\_I\_Msgs} + U \cdot L_{hash}$, $D\_Ovh_{CA} = N \cdot S \cdot (2 \cdot C_{verify} + C_{sign})$, and $D\_Ovh_{EVCP} = S \cdot (2 \cdot C_{sign} + C_{verify}) + U \cdot C_{hash}$.

The verifier's average query cost in CREV-I scheme is $Ovh_{Verifier\_Avg} = Pr_{rev} \cdot C_{hash} + Pr_{valid} \cdot (\frac{\ell_{CREV\_I}+1}{2} \cdot C_{hash} + C_{verify})$. The corresponding total daily costs due to query are: $D\_Bw_{EVCP-Ver}$ (for EVCP) $= q_{per\_cert\_daily} \cdot (L_{hash} + L_{CREV\_I\_Reply})$, $D\_Bw_{EVCP-Ver}$ (for Verifier) $= Q_{Ver\_daily} \cdot (L_{hash} + L_{CREV\_I\_Reply})$, $D\_Ovh_{CA} = 0$, $D\_Ovh_{EVCP} = 0$, and $D\_Ovh_{Verifier\_Avg} = Q_{Ver\_daily} \cdot Ovh_{Verifier\_Avg}$.

As in CRS, for a certificate with i already-released hash tokens ($0 \leq i \leq \ell_{CREV\_I}$), the CA only needs to store $(\ell_{CREV\_I} - i)$ hash-tokens together with N_0. Hence, the CA's stor-

age requirement is $Stor_{CA} = \sum_{i=1}^{\ell_{CREV\_I}} \frac{N}{\ell_{CREV\_I}} \cdot (i+1) \cdot L_{hash} = \frac{N}{\ell_{CREV\_I}} \cdot \frac{\ell_{CREV\_I}^2 + 3\ell_{CREV\_I}}{2} \cdot L_{hash}$. The storage needed in each EVCP is $Stor_{EVCP} = L_T + L_{CREV\_I\_Reply} + L_{hash}$, with L_T denoting the length of the CA's nonce for its replay protection (see Section 7.3.2).

For the CA's certificate creation, the (online) daily cost is $D\_Ovh_{CA} = \frac{N}{\beta} \cdot C_{sign}$. The timeliness guarantee of CREV-I scheme is δ minutes.

CREV-II

CREV-II establishes a short-term session of $t_{CREV-II}$ minutes between the CA and an EVCP. Throughout an established session, the CA sends an online status message (i.e. OCSP Response) every $d_{CA} = \delta$ minutes. In a day, each EVCP thus performs $S = \frac{M}{t_{CREV\_II}}$ session establishments, and receives $U = S \cdot \ell_{CA} = \frac{M}{\delta}$ OCSP Response messages. We denote the length of Status-Subscription Establishment Reply as $L_{CREV\_II\_Reply} =$ 550 bytes, and the length of all messages in a session establishment as $L_{CREV\_II\_Msgs} =$ 1,577bytes.

The total daily update costs (comprising both the session establishments and the OCSP Response updates) are: $D\_Bw_{CA-EVCP}$ (for CA) $= N \cdot (S \cdot L_{CREV\_II\_Msgs} + U \cdot L_{OCSP\_Response})$, $D\_Bw_{CA-EVCP}$ (for EVCP) $= S \cdot L_{CREV\_II\_Msgs} + U \cdot L_{OCSP\_Response}$, $D\_Ovh_{CA} = N \cdot S \cdot (2 \cdot C_{verify} + C_{sign}) + N \cdot U \cdot C_{sign}$, and $D\_Ovh_{EVCP} = S \cdot (2 \cdot C_{sign} + C_{verify}) + U \cdot C_{verify}$. With the range optimization technique, the worst case scenario for the required number of group-based status updates occurs when each revoked certificate happens to be in between two valid ones. As such, we require $Rev\_Entries$ OCSP Response messages. The average case requires $N_{range\_msgs} = \lceil \frac{Rev\_Entries}{2} \rceil$ messages. By denoting $L_{OCSP\_Response\_Range}$ as the length of OCSP Response with range specification, we have: D_Ovh_{CA} (with range optimization) $= N \cdot S \cdot (2 \cdot C_{verify} + C_{sign}) + N_{range\_msgs} \cdot U \cdot C_{sign}$, and $D\_Bw_{CA-EVCP}$ (for CA, with range optimization) $= N \cdot (S \cdot L_{CREV\_II\_Msgs} + U \cdot L_{OCSP\_Response\_Range})$.

The total daily query costs are as follows: $D\_Bw_{EVCP-Ver}$ (for EVCP) $= q_{per\_cert\_daily} \cdot L_{OCSP\_Response}$, $D\_Bw_{EVCP-Ver}$ (for Ver) $= Q_{Ver\_daily} \cdot L_{OCSP\_Response}$, $D\_Ovh_{CA} = 0$, $D\_Ovh_{EVCP} = 0$, and $D\_Ovh_{Ver} = Q_{Ver\_daily} \cdot C_{verify}$. With range optimization technique, we use $L_{OCSP\_Response\_Range}$ instead on $L_{OCSP\_Response}$ in $D\_Bw_{EVCP-Ver}$ calculations.

The storage requirements are: $Stor_{CA} = 0$ (without range optimization) or $Stor_{CA} = N_{range\_msgs} \cdot L_{OCSP\_Response\_Range}$ (with range optimization), and $Stor_{EVCP} = L_T + L_{OCSP\_Response}$. For the CA's certificate creation, the (online) cost is $D\_Ovh_{CA} = \frac{N}{\beta} \cdot C_{sign}$. The timeliness guarantee of CREV-II scheme is: $d_{CA} = \delta$ minutes.

7.4.4 Performance Evaluation

Evaluation Scenarios and Objectives

We conducted our evaluation on a certification scenario with 250, 000 certificates and 10% revocation rate. The other parameter values for two scenarios (with the timeliness guarantee of 1 and 10 minutes respectively) can be seen in Table 7.1.

Our scenarios also aim to capture an observation that the principals of EV certificates are queried by the verifiers more often than those of Regular ones. Therefore, while N_{EV} represents only 20% of N, the ratio of daily query between an EV and Regular certificate (μ) is 10,000, with $q_{per\_EV\_daily} = 1,000,000$ and $q_{per\_Reg\_daily} = 100$.

For the calculation of Pr_{rev} and Pr_{valid} (Eq. 7.23), we simply assume that 80% of the EV certificates follow $S_A(t)$ model and the other 20% observe $S_B(t)$. For Regular certificates, we set 60% of N_{Reg_cert} belong to $S_A(t)$ and 40% are of $S_B(t)$.

Our evaluations aim to compare CREV with other schemes under the following two settings: (i) when all certificates in a system are of EVC type; and (ii) when a certification system involves both the EV and Regular certificate types. For the second scenario, we assume the availability of a CMAE running CREV-II (with range optimization technique) that manages all the Regular certificates. It thus acts as a CMAE for the Regular certificate principals. When the CREV-II scheme with range-optimization is employed by this CMAE, its total daily update costs are (with $N_{range\_msgs} = \lceil \frac{Rev\_Entries}{2} \rceil$ as in Section 7.4.3): $D\_Bw_{CA-CMAE} = S \cdot L_{CREV\_II\_Msgs} + N_{range\_msgs} \cdot U \cdot L_{OCSP\_Response\_Range}$, $D\_Ovh_{CA} = S \cdot (2 \cdot C_{verify} + C_{sign}) + N_{range\_msgs} \cdot U \cdot C_{sign}$, and $D\_Ovh_{CMAE} = S \cdot (2 \cdot C_{sign} + C_{verify}) + N_{range\_msgs} \cdot U \cdot C_{verify}$. The total daily query costs are: $D\_Bw_{CMAE-Ver}$ (for CMAE) $= N_{Reg\_cert} \cdot q_{per\_Reg\_daily} \cdot L_{OCSP\_Response\_Range}$, and $D\_Ovh_{CMAE} = 0$. The CMAE's storage requirement is $Stor_{CMAE} = L_T + L_{OCSP\_Response\_Range}$. In calculating the verifier's overheads, we assume that a verifier's daily queries also follow the defined query ratio μ between the EV and the Regular certificates.

Evaluation Results

We set that all hash values are produced using SHA-1, and the signatures are created using RSA with 1024-bit modulus. Based on the Crypto++ 5.6.0 Benchmarks (`http://www.cryptopp.com/benchmarks.html`) on Intel Core-2 PC with 1.83 GHz CPU running Windows Vista 32-bit, we have the overheads of the basic cryptographic operations as shown in Table 7.1.

The results of realistic performance factors ($Rev\_Entries$, $q_{per\_cert\_daily}$, Pr_{rev}, and Pr_{valid}) under the specified scenarios are shown in Table 7.2. As can be seen, we have the $Rev\_Entries$ of 24,736.33 (for $\delta = 10$ mins) and 24,736.54 (for $\delta = 1$ min), both of which represent $\sim 98\%$ of the total revoked certificates. They are however significantly

higher from that in a simplified model based on an assumption that revoked entries are kept in the CRL for $\frac{\beta}{2}$ days [177, 84] (only 12,500 entries are recorded in the CRL). With $N = 250,000$, the difference in CRL data size becomes 477,204 bytes, which makes it an important consideration for the bandwidth-limited verifiers. We notice that, if we derive *Rev_Entries* by simply applying function $f(v)$ or $N(v)$ [91, 65], which uses a summation on the revocation PDF, we would have incorrect CRL entries of 102,493,910 (with $\delta =$ 10 minutes on 50,000 EV certificates with 10% revocation). This result particularly highlights the problem with summation-based calculation on PDF as in [91, 65] when δ is less than 1 day.

Overheads	$\delta = 10$ **minutes**			$\delta = 1$ **minute**		
Factors	**All Certs**	**EV Only**	**Reg Only**	**All Certs**	**EV Only**	**Reg Only**
Rev_Entries	24,736.33	4,947.27	19,789.06	24,736.54	4,947.31	19,789.23
CRLSize (KB)	942.48	188.80	754.07	942.48	188.80	754.07
Q_{daily}	5.002×10^{10}	5×10^{10}	2×10^{7}	5.002×10^{10}	5×10^{10}	2×10^{7}
$q_{per\_cert\_daily}$	200,080	1,000,000	100	200,080	1,000,000	100
Pr_{rev}	0.005563	0.005561	0.010835	0.005564	0.005562	0.010836
Pr_{valid}	0.994437	0.994439	0.989165	0.994436	0.994438	0.989164

Table 7.2: Calculation results of some performance factors under the specified scenarios.

Another important finding is that the probability of a verifier's query to be reported as valid is ∼98-99%. The probability is higher than 92% reported in [84] where it is simply assumed that $Pr_{valid} = N/(N + Rev\_Entries)$. The higher percentage is due to our more realistic exponential probability function $S(t)$, which models the "decaying" effect of queries on the revoked certificates over time. This result is particularly important for hash-chaining based schemes since now the verifier is *almost certain* to perform the repeated hash operations to prove the validity (instead of the revocation) of a queried certificate. Our CREV schemes, which keep the verifier's fast verification, thus become more significant particularly for the lightweight verifiers.

Lastly, we also can see that by applying Eq. 7.18, we have $Q_{EV\_daily} = 5 \times 10^{10}$. The number is 2,500 times larger than $Q_{Reg\_daily} = 2 \times 10^{7}$ despite the fact that $N_{Reg\_cert} = 4 \cdot N_{EV\_cert}$.

Table 7.3 shows the comparison of the performance overheads between CREV and other schemes when only EV certificates exist in the certification system ($N_{EV} = 50,000$) and $\delta = 10$ minute. Table 7.4 shows the comparison results on a certification system containing both the EV and the Regular certificates ($N = 250,000$) and $\delta = 1$ minute.

7.5 Discussion

We first discuss how CREV schemes compare with others on 50,000 EVC-only certificates. We highlight the results from Table 7.3.

Cost *(U=Update, Q=Query)*	**Unit**	**CRL**	**OCSP**	**CRS**	**CREV-I**	**CREV-II**
CA's Daily Costs:						
D_Ovh_{CA} (U+Q)	sec	0.21	$\mathbf{7.4 \times 10^7}$	0.21	162	10,818
D_Ovh_{CA} (Cert Creation)	sec	0.20	0.20	0.20	0.20	0.20
$Stor_{CA}$	MB	0.18	0	**25,063.13**	35.29	0
$D\_Bw_{CA-CMAE}$ (U)	MB	26.55	-	203.76	-	-
$D\_Bw_{CA-EVCP}$ (U)	MB	-	-	-	289.44	3,302.10
D_Bw_{CA-Ver} (Q)	MB	-	$\mathbf{2.19 \times 10^7}$	-	-	-
CMAE's Daily Costs:						
D_Ovh_{CMAE} (U+Q)	sec	0.01	-	0.01	-	-
$Stor_{CMAE}$	MB	0.18	-	1.47	-	-
$D\_Bw_{CA-CMAE}$ (U)	MB	26.55	-	203.76	-	-
$D\_Bw_{CMAE-Ver}$ (Q)	MB	$\mathbf{9.22 \times 10^9}$	-	$\mathbf{9.54 \times 10^5}$	-	-
EVCP's Daily Costs:						
D_Ovh_{EVCP} (U+Q)	sec	-	-	-	0.0061	0.0161
$Stor_{EVCP}$	MB	-	-	-	6.11×10^{-4}	4.53×10^{-4}
$D\_Bw_{CA-EVCP}$ (U)	MB	-	-	-	0.01	0.07
$D\_Bw_{EVCP-Ver}$ (Q)	MB	-	-	-	596.05	437.74
Verifier's Daily Costs (with 30 average daily queries):						
D_Ovh_{Ver} (Q)	sec	0.0021	0.0021	0.3136	0.0025	0.0021
D_Bw_{CA-Ver} (Q)	MB	-	0.0131	-	-	-
$D\_Bw_{CMAE-Ver}$ (Q)	MB	**5.53**	-	0.0006	-	-
$D\_Bw_{EVCP-Ver}$ (Q)	MB	-	-	-	0.0179	0.0131
Timeliness	min	10	≈ 0	10	10	10

Table 7.3: Performance comparison on EV-certificate only certification system (N_{EV}=50,000) with $\delta = 10$ minutes

.

Cost *(U=Update, Q=Query)*	**Unit**	**CRL**	**OCSP**	**CRS**	**CREV-I**	**CREV-II*** **+ CREV-II* CMAE for Reg Certs**
CA's Daily Costs:						
D_Ovh_{CA} (U+Q)	sec	2.13	$\mathbf{7.4 \times 10^7}$	2.13	21,736.24	27,008.83
D_Ovh_{CA} (Cert Creation)	sec	1.01	1.01	1.01	1.01	1.01
$Stor_{CA}$	MB	0.92	0	$\mathbf{1.25 \times 10^6}$	86.78	1.10
$D\_Bw_{CA-CMAE}$ (U)	MB	1,325.37	-	10,187.12	0.65	
$D\_Bw_{CA-EVCP}$ (U)	MB	-	-	-	1,981.74	32,599.26
D_Bw_{CA-Ver} (Q)	MB	-	$\mathbf{2.19 \times 10^7}$	-	-	-
CMAE's Daily Costs:						
D_Ovh_{CMAE} (U+Q)	sec	0.10	-	0.10	997.44	
$Stor_{CMAE}$	MB	0.92	-	7.36	0.0005	
$D\_Bw_{CA-CMAE}$ (U)	MB	1,325.37	-	10,187.12	0.65	
$D\_Bw_{CMAE-Ver}$ (Q)	MB	$\mathbf{4.604 \times 10^{10}}$	-	$\mathbf{9.54 \times 10^5}$	8,888.24	
EVCP's Daily Costs:						
D_Ovh_{EVCP} (U+Q)	sec	-	-	-	0.0248	0.1250
$Stor_{EVCP}$	MB	-	-	-	0.0006	0.0005
$D\_Bw_{CA-EVCP}$ (U)	MB	-	-	-	0.04	0.65
$D\_Bw_{EVCP-Ver}$ (Q)	MB	-	-	-	596.05	444.41
Verifier's Daily Costs (with 30 average daily queries):						
D_Ovh_{Ver} (Q)	sec	0.0021	0.0021	3.1361	0.0032	0.0021
D_Bw_{CA-Ver} (Q)	MB	-	0.0131	-	-	-
$D\_Bw_{CMAE-Ver}$ (Q)	MB	**27.61**	-	0.0006	1.33×10^{-6}	
$D\_Bw_{EVCP-Ver}$ (Q)	MB	-	-	-	0.0179	0.0133
Timeliness	min	1	≈ 0	1	1	1

Table 7.4: Performance comparison on a certification system containing both the EV and the Regular certificates ($N = 250,000$) with $\delta = 1$ minute. CREV-II* makes use of the range-optimization technique.

Given a CRL size of 188.80 KB for 50,000 principals, **the CRL scheme** incurs an enormous daily bandwidth of 9.22×10^9 MB between the CMAE and the verifier ($D\_Bw_{CMAE-Ver}$). Another real problem faced by CRL is that a verifier, who performs just 30 queries per day, needs to download 5.53 MB of daily CRL data which may not be reasonable for mobile devices.

In **OCSP**, a verifier receives only 0.031 MB of OCSP Responses daily. However, OCSP puts a large burden on the CA. Every day, the CA needs to deliver 2.19×10^7 MB of OCSP Response messages to all the verifiers, and spend 7.4×10^7 seconds of processing time to sign the Response messages.

CRS significantly reduces the bandwidth requirement from the CMAE to the verifiers. Each verifier also needs to access only ∼600 bytes of hash tokens daily. However, on average, a verifier needs to spend 10.45 ms processing *a single* hash operation (on an 1.83-GHz Intel Core-2 CPU). With the length of the hash chain of 52,559, the worst case for a single verification operation is 21.02 ms, which is much more expensive than verifying a digital signature (0.07 ms). As such, this overhead may be too costly for lightweight verifiers such as netbooks or mobile devices. Another problem is that CRS requires a huge storage on the CA (25,063.13 MB) when the CA stores all the hash chains (without amortization). Furthermore, the bandwidth requirement from the CMAE to all verifiers is very high (9.54×10^5 MB).

CREV-I offers an attractive trade-off between the CA's storage requirement and processing overheads while still maintaining a low verifier's computational (energy) and bandwidth requirements. The storage requirement on the CA is now down to 35.29 MB from 25,063 MB earlier in CRS/NOVOMODO. With still 0.08 ms (or 2.518 ms daily) of the verifier's average processing time on a query, the downsides of CREV-I are the CA's extra bandwidth and processing due to the establishments of sessions. The overall bandwidth of $D\_Bw_{CA-EVCP}$ is 289.44 MB/day, which is still reasonable. The CA's daily processing time becomes 162 s (∼2.7 minutes), which is significantly smaller compared to the 7.4×10^7 s or ∼2,056 hrs in OCSP.

CREV-II has a unique advantage that it delivers a standardized OCSP Response message. The advantage over OCSP is that the bandwidth requirement from the CA to all EVCP servers is only ∼3,302 MB/day, which is a very much smaller than 2.1×10^7 MB/day in OCSP. The CA's computational cost increases to 10,818 s or ∼3 hrs daily which is still quite affordable. CREV-II thus makes online status-based schemes much more viable to deploy.

We now discuss the scenario of a certification system comprising of EV and Regular certificates. The results are shown in Table 7.4. CREV schemes offer the following advantages over the other analyzed schemes.

The two CREV schemes only make small computation and bandwidth demands on

the verifier. With 30 queries daily, the highest verifier's daily bandwidth requirement is 0.0179 MB (in CREV-I), and verifier's daily computation overheads is 3.2 ms (in CREV-I). In a mobile device setting, the processing time is expected to be higher but still be reasonable.

Our CREV schemes also result in a reasonable processing and bandwidth requirements on the **EVCP server** and **the (proxy) CMAE**. Despite having to manage all the Regular certificates (80% of N), the CMAE running CREV-II (with the range-optimization technique) requires only ∼8,888 MB of $D\_Bw_{CMAE-Ver}$, which is very much smaller than 4.604×10^{10} in CRL and 9.54×10^5 in CRS. This is due to our scenario's assumptions in which most of the CSI queries are issued on the online-service providing principals holding EV certificates. Our use of the EVCP servers to deal with their respective verifiers works well to distribute the workload.

For the **CA's overheads**, CREV schemes manage to avoid high CA's processing and bandwidth overheads suffered in OCSP with its centralized CA processing. Moreover, our use of session-based hash chains enables us to significantly reduce the CA's requirement on storing the hash chains. CREV-I only requires 86.78 MB of storage on the CA, which is a significant reduction from 1.25×10^6 MB in CRS.

In summary, both CREV-I and CREV-II offer a good trade-off between the costs incurred on CA, CMAE and EVCP, while maintaining fairly lightweight requirements on the verifier. Compared with other examined schemes, these two CREV schemes are more viable for deployment when a near real-time timeliness guarantee is needed. CREV-II incurs a higher overhead on the CA's processing compared to CREV-I due to a higher volume of signing operations. However, with the range-optimization technique, this is still reasonable with a daily computation overhead of ∼7.5 hrs.

The main new requirement for our CREV schemes is that an EVCP now needs to be available and provide a reliable CSI access service. However, given the fact that an EVCP needs to be up and running to provide its own online services, this requirement is theefore not an issue. With a present trend of using *Content Delivery Network (CDN)* which are commercially available from various parties (e.g. Akamai Technologies and Amazon CloudFront), we envision that the EVCP servers are able to provide the uptime requirement either by themselves or through the use of third party providers. Lastly, we also note that the CREV schemes can be deployed in conjunction with other standard revocation schemes such as CRL or OCSP, thus ameliorating the potential workload bottlenecks faced in the existing revocation schemes.

7.6 Conclusion

We have presented two lightweight, practical and inherently-distributed certificate revocation schemes, called CREV, based on the recently available Extended-Validation Certificate infrastructure. Based on the analysis using our realistic performance analysis framework, we have shown the practicality of our schemes when compared to several existing schemes under the scenarios of revocation timeliness guarantees between 1 and 10 minutes. The CREV-I and CREV-II schemes offer a good balance of incurred costs on the involved entities, while maintaining lightweight requirements on the verifier. With the range-optimization technique, even with a timeliness guarantee of 1 minute, CREV-II can keep the CA's computational overheads manageable. Given their practical setting-implementability and lightweight requirements, the proposed CREV schemes are thus very suitable for certificate revocation in numerous real-world scenarios. With their near real-time timeliness guarantee, the schemes can offer appealing alternative solutions to the problem of providing a timely yet lightweight revocation infrastructure, including to mobile devices where computation, bandwidth and energy usage need to be kept small.

Chapter 8

Extending BAN Logic for Reasoning with PKI-based Protocols

It is common nowadays for a running program to interact with external entities over a public insecure network like the Internet. In order to secure the communication and the access to provided services, public-key cryptographic operations are often employed. PKI-based protocols, including those used in securing program distribution and certificate revocation management of the PPLC, must be shown to be secure. Designing a correct protocol specification is however well recognized as a difficult task. Hence, a formal analysis is necessary to establish the security of the protocols. Despite the availability of numerous formal techniques, there is a challenge to devise one which can be handily utilized by many protocol designers to verify real-world protocols.

Among various authentication logics, BAN Logic [25, 26] (see also Section 2.6 for its background information) is one of the best known and most widely used techniques [98, 133, 99]. This may well explain the constant appearance of publications applying BAN Logic to protocols and security systems even till now [4, 23, 143, 175, 34, 32], with application domains as diverse as wireless network [175], mobile communication [32] and voting system [143]. BAN Logic however does not properly deal with the detailed aspects of PKI-based authentication, such as certificate processing, as commonly practiced nowadays. This is arguably because PKI was not well established yet when the logic was designed. The situation now is however very different since PKI becomes common, and is an infrastructure on which many real-world protocols critically rely. Given the ubiquity of PKI-based protocols and their importance to host security, there is a need to update BAN Logic to better reason with PKI-based protocols, yet remain practical to use.

This chapter explains our extension of BAN Logic, published in [147], which allows for a more concise reasoning with PKI-based protocols. In our work, we begin with

the starting point of retaining the popularity of BAN Logic among protocol designers. Our extension is along the lines of the work by Gaarder and Snekkenes [48] but better captures the current aspects of PKI. We address various limitations of [48] in capturing many important concepts and practices of the modern PKI usage. We also apply our logic to verify the session establishment protocol in CREV scheme proposed in Chapter 7.

The remainder of this chapter is organized as follows. We first survey the related works in Section 8.1. We then give a brief review of the previous extension of BAN Logic by Gaarder and Snekkenes in Section 8.2. Section 8.3 presents our new extended PKI-based BAN Logic, whereas Section 8.4 gives insight on its usage in preventing flawed protocol designs. An example of the application of our logic is given in Section 8.5. Section 8.6 discusses the results, and Section 8.7 finally concludes this chapter.

8.1 Related Work

There exist various works in the literature, such as [4, 48, 144, 95, 77, 64, 149], which apply formal methods to PKI. We briefly survey ones which extend authentication logics, particularly BAN Logic, to deal with public-key authentication. Our focus of comparison will be on certificate processing formalism, time durations, and rules on messages encrypted (signed) using public (private) keys.

As pointed out by many researchers, such as in [14], the original BAN Logic is known to have limitations in describing "serverless protocols". In PKI setting, the limitations have to do with accepting the validity of a certificate. This may happen since the only way of promoting "once said" ($\mid\!\sim$) to "believe" ($\mid\!\equiv$) is by use of the freshness property of a statement, which is typically in the form of a nonce or a timestamp. In a serverless protocol, such freshness guarantee however cannot be provided, because the server is not necessarily available at the time of communication. To work around this problem, the original BAN Logic choose to ignore the initial handling of certificates by assuming that they have been previously distributed, checked, and accepted as valid. Aziz and Diffie, who applied BAN Logic in [14], alternatively assume a certificate to always be fresh. Hence, the required belief statements on certificate contents can somehow be derived.

In their work on formal verification of CCITT X.509 protocol [48], Gaarder and Snekkenes argued that important aspects of public-key authentication are lost when BAN Logic is used for protocol verification. To amend this deficiency, they proposed enhancements to BAN Logic that take certificate checking into account as an integral part of the reasoning process. The extension defines the notion of *duration* to capture some time-related aspects. A principal can therefore claim that a formulae is, was, or will be good in a time interval. In [144], Stubblebine and Wright however argued that the assumptions used are too restrictive for reasoning about long-lived security

associations. Additionally, there exist issues on synchronization and synchronization bounds. Nevertheless, the simplicity of the logic proposed in [48], while improving the ability to reason with PKI-based protocols, is appealing. Our work here focuses on reworking the logic to be more accurately in line with the current PKI practice.

The work of Syverson [149] also adds the element of time to a logic of authentication. It incorporates a temporal formalism into a semantic model of BAN Logic developed in [2], using temporal notions of "all points in the run prior to the current one" and "at some point in the run prior to the current one." Here in our work, we adopt the duration model of [48] which is relatively easier to use, yet enables the analysis of subtle relationships in PKI-based protocols.

Stubblebine and Wright [144] also propose a logic extension for dealing with PKI. The logic supports the concept of synchronization, revocation and recency. In pursuing more expressiveness, it however becomes far more complex than the original BAN Logic. We view that the complexity is a drawback which makes it less likely to be used in practice.

Parts of the extension logic appeared earlier in the author's Master's thesis [145], which described the basic ideas of the extension logic. These basic ideas were later extended into a published work [147] and explained in this chapter. The logic formulation in [145] only covers the basic aspects and is not full fledged. More specifically, it does not cover: the stated sender (constructs 8.15 and 8.16—see later), the related Message-meaning Rule for public-key encrypted message (Rule 8.17), as well as the analysis of encrypted signed message (Rule 8.18) and signed encrypted message (Rules 8.19– 8.22). In addition, it also lacks a comparative analysis of the logic with other works that incorporates time into BAN Logic as discussed earlier in this section. Also missing in [145] is the analysis of how the extension can help prevent BAN Logic's "imprecise proof" when applied to protocols (elaborated later in Section 8.4). The reported results in this chapter are based on the work [147] which was published during the author's Ph.D candidature.

8.2 The Extension by Gaarder-Snekkenes

The extension logic of Gaarder and Snekkenes [48], which we call here **GS-BAN**, keeps BAN Logic's secret-key aspects intact for easy application. Our proposed extension takes the same approach. Appendix C lists all the secret-key rules of (the original) BAN Logic which are relevant to our extension in this chapter. Below, we summarize the extension logic of GS-BAN, and pinpoint some problems with it.

8.2.1 GS-BAN Extension Summary

Constructs for Public-Key Formalism

The following logic constructs were defined for public-key authentication in [48] (we use

some slight notational modifications on keys below):

- $\wp\kappa\,(P, K_P)$: Principal P has associated a good public key K_P ;
- $\Pi(K_P^{-1})$: Principal P has a good private key K_P^{-1} ;
- $\sigma(X, K_P^{-1})$: X signed with P's private key K_P^{-1} ;
- $\{X\}_{K_P}$: X encrypted under P's public key K_P.

In GS-BAN extension, it is assumed that a digital signature is always in *appendix mode*. (We also adopt the same assumption here.) Hence, the signature construct $\sigma(X, K_P^{-1})$ is actually a contracted form of the following:

$$\sigma(X, K_P^{-1}) = X, Sign(K_P^{-1}, hash(X)) \tag{8.1}$$

where: $Sign()$ is a signature function, and $hash()$ is a hash construction function.

Certificate and Its Idealization

As mentioned earlier, one of the extension's main contributions is the idealization of a certificate. It adopts the certificate based on the X.509 standard [68], whose basic structure can be described as follows:[1]

$$Cert_P = \sigma((N, I, \delta^P, P, K_P, A), K_I^{-1}) \tag{8.2}$$

where:

- N : unique serial number of the certificate;
- I : name of the issuer;
- δ^P : validity period of the certificate, which consists of t_1^P (not before) and t_2^P (not after);
- P : the distinguished name of the principal;
- K_P : the certified public key of P;
- A : identifier of the signature algorithm employed;
- K_I^{-1} : I's private key which is used to sign the certificate.

Based on the certificate structure, GS-BAN gave a certificate the following idealization:

$$Cert_P = \sigma((\Theta(t_1^P, t_2^P), \wp\kappa\,(P, K_P)), K_I^{-1}). \tag{8.3}$$

In its idealized form, a certificate contains a newly defined *validity period* construct, which is also called "duration-stamp" in [48]. The construct $(\Theta(t_1, t_2), X)$ was specifically introduced to say that "statement X holds in the time interval $[t_1, t_2]$". A message might thus be tagged with the duration-stamp denoting the time interval during which its creator claims the message is good. In a certificate, the tagged message represents the

[1] To simplify the formalism, GS-BAN extension considers the certification path to be of length 1. We also take similar approach here.

certificate statement $C^P = \wp\kappa\,(P, K_P)$ in GS-BAN. Hence, the issuer I who uttered the duration-stamped certificate in (8.3) claims that $\wp\kappa\,(P, K_P)$ is good in the time interval of $[t_1^P, t_2^P]$.

Inference Rules

To reason with the certificate and public-key constructs, GS-BAN introduced the following inference rules:

- The Message-meaning (for public-key) Rule:

$$\frac{P \models \wp\kappa\,(Q, K_Q)\,, P \models \Pi(K_Q^{-1}), P \triangleleft \sigma(X, K_Q^{-1})}{P \models Q \mid\sim X} \tag{8.4}$$

- The See signed-message Rule:

$$\frac{P \triangleleft \sigma(X, K_Q^{-1})}{P \triangleleft X} \tag{8.5}$$

- The Certificate duration-stamp Rule:

$$\frac{P \models Q \mid\sim (\Theta(t_1^R, t_2^R), C^R), P \models Q \models \Delta(t_1^R, t_2^R)}{P \models Q \models C^R} \tag{8.6}$$

  In the above rule, Q acts as the issuer of a certificate for R, whose certificate statement is denoted by C^R. The rule thus provides a way of promoting *once said* to *believe* about a certificate statement. For this purpose, P needs to believe that Q holds a belief on **"good time interval"** t_1^R and t_2^R (denoted by $\Delta(t_1^R, t_2^R)$). In other words, P must believe that the validity period of $[t_1^R, t_2^R]$ still holds according to the current time in Q.

Message-Recipient Construct

GS-BAN also introduced a notion of "intended recipient" of a message. Its authors argued that such assurance needs to be explicitly stated as one of the goals in their X.509 protocol formalism. Moreover, they also claimed that the original BAN Logic provides no means of expressing it in the language available. Therefore, the construct "$\Re(X, P)$" was introduced to say that "P is the intended recipient of message X".

The construct is defined to appear in the following form:

$$P \rightarrow Q : X, \Re(X, Q) \tag{8.7}$$

which should be interpreted as "P sends to Q a particular message X together with a statement telling that Q is the intended recipient of X". This effort clearly represents a step forward in capturing the notion of *explicit principal naming* [1]. In doing so, however, GS-BAN requires the receiver to hold an assumption about the sender's jurisdiction over the message recipient statement (i.e. $Q \models P \Rightarrow \Re(X, Q)$), an approach which we will examine more closely below.

8.2.2 Problems and Limitations

Despite its improvements on BAN Logic, there exist some key aspects of GS-BAN extension [48] which we find rather unsettling:

- **Assumption on** $\Pi(K_Q^{-1})$: In verifying a protocol using GS-BAN, a (supplied) assumption needs to be added to a principal that he/she believes the goodness of the private key of *another principal* involved. That is, a formula such as "$P \models \Pi(K_Q^{-1})$" needs to be *vacuously held* at the start of the verification process. Such stipulation is needed so that GS-BAN can process the Message-meaning Rule (8.4). We however view the inclusion of such assumption is an unsound practice. Such a formula should never be supplied as an assumption, but must rather be logically derived from a certificate reasoning within the logic. The only assumption needed is the goodness of public and private keys of the CA in addition to the principal's own key pair. A principal then should be able to derive the goodness of the other principal's keys from the certificates issued by the CA. Furthermore, if a key is no longer considered good, revocation mechanisms like CRL [36] should be incorporated in the logic.

- **Assumption for message recipient construct**: One of the consequences of the Message-recipient construct in GS-BAN is that it requires a statement $\Re(X, P)$ to be made as a protocol goal for every message with the intended recipient construct. In our opinion, such statement does not need to be stipulated as a formalism goal since we are mainly interested in deriving goals about the goodness of the generated session key(s). In our view, the intended recipient construct should instead be incorporated as a part of logic rules. Moreover, as mentioned above, the construct requires the receiver to hold an assumption about the sender's jurisdiction over the message recipient statement (i.e. $Q \models P \mapsto \Re(X, Q)$). In our new extension, by integrating the intended-recipient requirement into the new Message-meaning Rule, we manage to eliminate the need for sender's jurisdiction, thus simplifying the reasoning on message processing.

- **Omission of certificate revocation process**: GS-BAN apparently has chosen to ignore the incorporation of certificate revocation issue into its logic. Instead, it assumes that each principal always maintains the goodness of its private key. As a result, we need to believe that a certificate, once issued, is *always good* for the time-interval specified in its validity period. Considering the importance of certificate revocation in public-key authentication, its omission in the formalization represents a limitation of the extension.

8.3 MPKI-BAN: Extending BAN Logic to Deal with PKI

Motivated by the aforementioned drawbacks of GS-BAN extension [48], we propose the following extensions to BAN Logic which addresses various limitations of GS-BAN and also provides other substantial contributions. We call our new extension **MPKI-BAN**, as it is aimed at dealing with modern PKI-based protocols.

Our approach of presenting the results in this chapter focuses on pragmatism. We focus on the logic definition and its usage application while leaving theoretical analysis of the logic, e.g. the logic's soundness and completeness with respect to some well-defined semantics, as a separate treatment beyond our work's scope. Our goal here is to expound an up-to-date yet accessible authentication logic which can be handily used by protocol designers who may not be experts in authentication logics or formal methods.

8.3.1 Revised Idealized Certificate

In MPKI-BAN, we idealize a certificate's statement as follows:[2]

$$Cert_P = \sigma(\Theta'(t_1^P, t_2^P, I, (\wp\kappa\,(P, K_P)\,, \Pi(K_P^{-1}))), K_I^{-1}) \tag{8.8}$$

To capture the need to ensure that a certificate is not revoked prior to believing it, we define "*validity period (duration-stamp) with revocation*" construct $\Theta'(t_1, t_2, I, X')$. This construct states that "(certificate) statement X' holds in the time interval $[t_1, t_2]$, *provided that* it is not revoked by the issuer I". We will redefine the certificate validation rule using this construct later in Section 8.3.5. We may keep the original validity period construct of GS-BAN ($\Theta(t_1, t_2), X$), now restated as $\Theta(t_1, t_2, X)$, to represent a validity period *without revocation*. Thus, the statement X will always be good in $[t_1, t_2]$ without any need to revalidate it with the sender. For a certificate in the X.509-based PKI framework, however, we need to use the newly defined function $\Theta'()$ rather than $\Theta()$.

As can be seen, we now include $\Pi(K_P^{-1})$ into the idealized certificate's statement in contrast with the previous definition in (8.3). Hence, a valid certificate now assures that *both public and private keys* of a principal are good. With this modification, we thus eliminate the need to have an assumption about the goodness of the other principal's private key as in GS-BAN. This modification is crucial as it helps formalize the close relationship between the certificate validity and the goodness a private key. The rule makes it clear that a belief on $\Pi(K_P^{-1})$ should be derived from a valid P's certificate. Consequently, should the private key of a principal ever be compromised, the principal must notify and ask the CA to revoke his/her certificate, a requirement that is consistent with current PKI practices.

[2]Note that the *complete* idealized certificate definition will include message-recipient construct as defined in (8.12).

8.3.2 New Use of Message-Recipient

Different from GS-BAN, the message-recipient construct is defined in our extension to be part of signature construct (σ), which now appears in the following form:

$$\sigma(\Re(X, P), K_Q^{-1}). \tag{8.9}$$

Note that now we make statement $\Re(X, P)$ as a stronger statement than X, as it implies a statement X together with its recipient tag for P.

Unlike the use of the message-recipient construct in GS-BAN, we incorporate it into our new Message-meaning Rule (defined below) as *one of its premises*. A principal thus needs to ensure the existence of a valid recipient tag in the signed message in order to proceed with the rule. With this, we also manage to eliminate the requirement of stating a message-recipient as a verification goal, as well as the requirement of introducing an assumption about the sender's jurisdiction over the message-recipient as in GS-BAN.

8.3.3 New Message-Meaning Rule for Private-Key Signed Messages

In symmetric-key based authentication, the intended recipient of a message can usually be inferred from the shared secret-key employed in the encryption or the generation of Message Authentication Code (MAC). Unfortunately, this does not apply in private-key signed messages where the same private key is used to sign messages regardless of their intended recipient. Hence, in public-key authentication, there is a greater need to follow the "naming principle" [1].[3] Surprisingly, the same mistake due to disregarding this principle seems to be made time after time in many protocols (see e.g. [89]). Given this concern, we redefine the **Message-meaning for (private-key) signed message** Rule as follows:

$$\frac{P \models \wp\kappa\,(Q, K_Q)\,, P \models \Pi(K_Q^{-1}), P \triangleleft \sigma(\Re(X, P), K_Q^{-1})}{P \models Q \mid\sim X}. \tag{8.10}$$

With (8.10), we thus integrate the message-recipient construct into the Message-meaning Rule. Our requirement for the third premise above is strongly motivated by the works of Meadows [98] and Boyd and Mathuria [24], which still managed to uncover a loophole in the Aziz-Diffie protocol [14] despite the application of BAN Logic to the proposed protocol. We will examine this later in Section 8.4 where we show how the new rule and some cautionary note on the application of BAN Logic can together help pinpoint the problem with the protocol and its imprecise proof.

[3] It is possible to argue that in a few situations, due to privacy considerations, protocol designers might aim to withhold identity information as long as possible, thus conflicting with the naming principle. We however focus here on general situations where ensuring secure authentication takes precedence over privacy concerns.

8.3.4 All-Recipient See Rule

Having defined the new Message-meaning Rule in (8.10) above, we further note that it is possible for a message to be actually intended for *all* principals in the protocol. A good example is a certificate, which is meant to be accepted by any principal as long as he/she trusts the issuer. We thus define a special principal name "*all*", and define the following rule:

$$\frac{P \triangleleft \sigma(\Re(X, all), K_Q^{-1})}{P \triangleleft \sigma(\Re(X, P), K_Q^{-1})}. \tag{8.11}$$

8.3.5 Certificate and Certificate-Validation Rule

In line with the use of *all recipient* definition above, a certificate is now to be idealized in MPKI-BAN as follows:

$$Cert_P = \sigma(\Re(\Theta'(t_1^P, t_2^P, I, (\wp\kappa\,(P, K_P)\,, \Pi(K_P^{-1}))), all), K_I^{-1}). \tag{8.12}$$

This certificate definition now correctly includes the message-recipient construct, and subsumes the previous interim definition given in (8.8).

To derive a belief on a certificate, this **Certificate-Validation Rule** is used:

$$\frac{P \models Q \mid\sim \Theta'(t_1^R, t_2^R, Q, C^R), P \models Q \models \Delta(t_1^R, t_2^R), P \models Q \models \Phi(C^R)}{P \models Q \models C^R}. \tag{8.13}$$

Here, Q acts as the certificate issuer. C^R denotes a certificate statement, consisting of $\wp\kappa\,(P, K_P)$ and $\Pi(K_P^{-1})$. This rule thus supersedes Rule (8.6) previously defined in GS-BAN. The third premise (e.g. $P \models Q \models \Phi(C^R)$) emphasizes the need for "*certificate revalidation step*" with the issuer Q as the principal stated in $\Theta'()$ construct. P must ensure this premise by checking that Q still believes that the uttered certificate statement remains valid ($Q \models \Phi(C^R)$).[4] In the CRL model, this step is done by checking the absence of the certificate in question in Q's recent CRL.

In the rule above, we note that the resulting belief statement (C^R) can be argued as an "unstable" statement [26]. That is, the statement is valid only at the time of validation but not necessarily thereafter, as the corresponding certificate might be revoked at some point of time in the future. A more elaborate logic might include a more general time-related reasoning as exemplified by the work [144]. However, in order to keep our extension simple, we avoid doing so. Instead, we limit such statements to beliefs on *the goodness of public and private key* of a principal, and not on actual statements communicated among the principals in the protocol. In fact, both BAN Logic [26] and GS-BAN [48] implicitly make a similar simplification too, because a secret key, in practice, will eventually cease to be valid due to its expiry or a possible security breach.

[4]Note that although we put C^R as the parameter of $\Phi()$, in practice the matching is done based on the unique certificate's serial number N.

8.3.6 Duration-Stamp (without Revocation) Validation

Having mentioned that we can keep a validity period without a revocation construct (see Section 8.3.1), we can define the following **Duration-stamp (without revocation) Validation Rule**:

$$\frac{P \models Q \mid\sim \Theta(t_1, t_2, X), P \models Q \models \Delta(t_1, t_2)}{P \models Q \models X}. \tag{8.14}$$

This rule thus promotes *once said* to *believe* on a duration-stamped (without revocation) statement. Here, P needs to ensure that the validity period of $[t_1, t_2]$ still holds according to the current time in Q. This rule is similar to the certificate duration-stamp Rule (8.6) in GS-BAN. For validating a certificate, MPKI-BAN however makes use of the more comprehensive Rule (8.13), which also includes a revocation checking step.

8.3.7 Message-Sender Construct

Realizing the importance of naming principle, we take another step to define a *message-sender* construct in MPKI-BAN. It aims to introduce the notion of a "stated sender" of a message. The message-sender construct $\mathcal{S}(X, Q)$ is defined to specifically say that "message X together with Q as the stated sender of the message". It can appear in the following two forms:

- Stated-sender within the encrypted message:
  This message-sender construct appears in the following form:
  $$\{\mathcal{S}(X, Q)\}_{K_P}. \tag{8.15}$$
  It occurs when the encryption is employed with an additional function of authentication.[5] Principal P, who receives the message from Q, ensures this construct by first decrypting the message, and then verifying that it contains Q as the sender ID together with correctly formatted message X. We define the corresponding Message-meaning for public-key encryption Rule in Section 8.3.8.
- Stated-sender outside of the encrypted message:
  This form appears as follows:
  $$\mathcal{S}(\{X\}_{K_P}, Q). \tag{8.16}$$
  It takes place when the encrypted message ($\{X\}_{K_P}$) comes in a message signed by Q. We discuss the usage of this construct and its related rule in Section 8.3.10.

8.3.8 Message-Meaning for (Public-Key) Encrypted Message Rule

When receiving a private-key signed message, it is important to ensure the intended recipient of the message. This requirement is captured by the message-recipient construct.

[5] It is thus important to make clear of the role of encryption in a protocol specification (see e.g. [1]).

When dealing with a message encrypted with the public key of a recipient, it is the *identity of the sender* that matters.

We consider the two following commonly applicable scenarios where a public-key encrypted message portion appears in a message. The first is when the encryption is employed as a form of authentication in addition to providing confidentiality. This occurs with the whole message is encrypted. The second scenario occurs when an encrypted message portion is contained in a signed message. We deal with the first scenario here, whereas the second scenario is analyzed in Section 8.3.10 below.

As we require a stated sender assurance on a public-key encrypted message, the decrypted message must contain the sender's identity in a pre-defined location in the message. The first type of the message-sender construct, $\{\mathcal{S}(X,Q)\}_{K_P}$ (8.15), applies in this scenario. That is, the ID of the sender is within the encrypted message. We thus define the following **Message-meaning for (public-key) encrypted message Rule**:

$$\frac{P \models \wp\kappa\left(P, K_P\right), P \models \Pi(K_P^{-1}), P \triangleleft \{\mathcal{S}(X,Q)\}_{K_P}}{P \models Q \mid\sim X} \tag{8.17}$$

We later show in Section 8.4 how this rule could have helped deal with a loophole in Needham-Schroeder Public-Key protocol, which is vulnerable to attacks [89].

Note that it is possible that the decrypted message X is actually a signed message. This case is considered further below (Section 8.3.9).

8.3.9 Additional Message-Meaning Rule for Encrypted Signed Message

The Message-meaning Rule above (8.17) specifies how MPKI-BAN Logic deals with a public-key encrypted message. In the case where the decrypted message is actually a signed message, we thus encounter an *encrypted signed message*, i.e. a signed message which is sent encrypted.

For the already-decrypted message, we define the following **Message-meaning for signed message Rule (after decryption)**:

$$\frac{P \models \wp\kappa\left(Q, K_Q\right), P \models \Pi(K_Q^{-1}), P \mid\sim \sigma(\Re(X,P), K_Q^{-1})}{P \models Q \mid\sim X}. \tag{8.18}$$

8.3.10 Rule for Signed Encrypted Message

Here we address another subtle issue, namely the reasoning of *signed encrypted message*, that is a signed message that contains encrypted message portion(s). Although the commonly suggested practice is to perform signing before encrypting [6], it is not uncommon to find the reverse in some published protocols.[6]

[6]Syverson [150] analyzed this ordering principle as suggested in [6], and questioned its applicability by putting forward examples when the principles are not applicable. Nevertheless, Syverson agreed that such principle is useful as a general guideline, although it should be critically used by the protocol designers.

Hence, after the message has been validated using the Message-meaning for signed message Rule (8.10), we need to specify how to process the encrypted message portion. Here, there are two possibility of the message-sender constructs as in (8.15) and (8.16):

- Stated-sender within the encrypted message ($\{\mathcal{S}(X,Q)\}_{K_P}$):

  Here, the sender ID is part of the encrypted message. We define the following **Message-meaning for encrypted message portion Rule (with included stated-sender information)**:

$$\frac{P \models \wp\kappa\left(P, K_P\right), P \models \Pi(K_P^{-1}), P \models Q \mid\sim \{\mathcal{S}(X,Q)\}_{K_P}}{P \models Q \mid\sim X}. \tag{8.19}$$

- Stated-sender outside of the encrypted message ($\mathcal{S}(\{X\}_{K_P}, Q)$):

  For this case, it is *insufficient* to simply take the presence of a sender ID in the (correctly signed) clear-text message portion. This is due to the "proof of authorship" problem of the encrypted message (see [6]). The problem is that we really cannot infer that the sender *actually knows* about X. To resolve this problem, we require the sender *to prove his knowledge of* X. On way to fulfill this requirement is by requiring the sender to generate the signature over the message, including X. Thus, X is calculated in the signature generation although its (clear-text) value does not appear in the signed message. One example protocol performing this is a proposed enhancement in [24] to fix the broken Aziz-Diffie protocol [14], which is analyzed later in Section 8.4.2. In this way, the message-sender construct $\mathcal{S}(\{X\}_{K_P}, Q)$ can be satisfied. In validating the signature, the receiver thus needs to first decrypt $\{X\}_{K_M}$ into X, and check that the signature is correctly constructed using X as a part of the input message.

  We can define the following **Message-meaning for encrypted message portion Rule (with external stated-sender information)**:

$$\frac{P \models \wp\kappa\left(P, K_P\right), P \models \Pi(K_P^{-1}), P \models Q \mid\sim \mathcal{S}(\{X\}_{K_P}, Q)}{P \models Q \mid\sim X}. \tag{8.20}$$

Notice that without the message-sender assurance, the receiver P can only derive $P \models Q \mid\sim \{X\}_{K_P}$, and not $P \models Q \mid\sim X$. This difference is crucial, and disregarding it can lead to erroneous proof generations using the original BAN Logic.

In order for the consequences of (8.19) and (8.20) to be derived further using the Nonce-verification Rule (R_2 in Appendix C), X needs to come with a freshness assurance so that $P \models \sharp(X)$ can be derived. There are two possible sources for such a freshness:

- The freshness assurance is inside the encrypted message:

  Here, statement X contains freshness assurance Y, which can be a nonce or a timestamp. As such, $P \models \sharp(Y)$ is assumed. Given the Freshness-extension Rule (R_{10} in Appendix C), we therefore can derive the required $P \models \sharp(X)$.

- The freshness assurance is not within the encrypted message:
  In this case, the freshness assurance must come in the clear-text portion outside of the encrypted portion ($\{X\}_{K_P}$) in the signed message. To emphasize the need for a freshness assurance in this case, the third premise in Rules (8.19) and (8.20) must be applied on both the encrypted message portion *and* a "fresh" statement. Hence, for completeness, we also define the following two rules:
  - **Message-meaning for encrypted message portion Rule (with encrypted stated-sender information, and external freshness assurance)**:
  $$\frac{P \models \wp\kappa\,(P, K_P)\,, P \models \Pi(K_P^{-1}), P \models Q \mid\sim \{\mathcal{S}(X, Q)\}_{K_P}, Y}{P \models Q \mid\sim X, Y}. \tag{8.21}$$
  - **Message-meaning for encrypted message portion Rule (with external stated-sender information and external freshness assurance)**:
  $$\frac{P \models \wp\kappa\,(P, K_P)\,, P \models \Pi(K_P^{-1}), P \models Q \mid\sim \mathcal{S}(\{X\}_{K_P}, Q), Y}{P \models Q \mid\sim X, Y}. \tag{8.22}$$

  Statement Y in the two rules above is the statement in the signed message which contains the freshness assurance, i.e. a timestamp or a nonce. Since both the encrypted message portion and Y come from the same single (signed) message, the required premises can be directly derived using the Said And-elimination Rule (R_7 in Appendix C). Given that $P \models \sharp(Y)$, we can derive $P \models \sharp(X, Y)$ using the Freshness extension Rule (R_{10} in Appendix C). As a result, the belief on X can be derived.

8.3.11 Redefined Message-Meaning Rule for Secret-Key

For completeness, we redefine here a new construct for message authentication using secret-key based MAC:

$$\mu(X, K_{PQ}) = X, H(K_{PQ}, X). \tag{8.23}$$

Similar to signature construction in (8.1), MAC generation is performed by applying a chosen keyed hash-construction function $H()$ to message X using K_{PQ}, and then appending the resulting hash image to X. The related Message meaning (for secret-key using MAC) Rule is defined as follows:

$$\frac{P \models Q \overset{K_{PQ}}{\longleftrightarrow} P, P \triangleleft \mu(X, K_{PQ})}{P \models Q \mid\sim X}. \tag{8.24}$$

Here we omit the requirements for the intended-recipient and stated-sender by assuming that the shared secret key is good and message X is unambiguously formatted.

8.3.12 Additional Rules for See Operator

- The **See hashed-message** Rule:

$$\frac{P \triangleleft \mu(X, K_{PQ})}{P \triangleleft X} \tag{8.25}$$

- The **See signed-message** Rule:

$$\frac{P \triangleleft \sigma(X, K_Q^{-1})}{P \triangleleft X} \tag{8.26}$$

- The **See recipient-tagged-message** Rule:

$$\frac{P \triangleleft \Re(X, P)}{P \triangleleft X} \tag{8.27}$$

For easy reference, all the newly added rules in MPKI-BAN are listed in Appendix D.

8.4 Using MPKI-BAN Logic

We now show how our MPKI-BAN could have helped prevent problems in vulnerable published protocols.

8.4.1 Needham-Schroeder Public-Key Authentication Protocol

In [89], Lowe published an attack on Needham-Schroeder public-key protocol. The protocol proceeds as follows (Notation $\{X\}_K$ below denotes an encryption operation using key K. When K is a private key, it represents a signing operation.):

1. $A \rightarrow S : A, B$
2. $S \rightarrow A : \{K_B, B\}_{K_S^{-1}}$
3. $A \rightarrow B : \{N_A, A\}_{K_B}$
4. $B \rightarrow S : B, A$
5. $S \rightarrow B : \{K_A, A\}_{K_S^{-1}}$
6. $B \rightarrow A : \{N_A, N_B\}_{K_A}$
7. $A \rightarrow B : \{N_A\}_{K_B}$.

Despite the assumption that each principal has each other's public key correctly, Lowe managed to find an attack on the protocol. The problem with the protocol has to do with the encryption using public key of the recipient without clear identity of the sender. In order to prevent potential attacks, Lowe proposed the modification of message 6 into:
$6'$. $B \rightarrow A : \{\boldsymbol{B}, N_A, N_B\}_{K_A}$.

In MPKI-BAN, with our new Message-meaning (for public-key encrypted message) Rule, the derivation of the incorrect beliefs would be impossible. This is because the required statement $\{\mathcal{S}((N_A, N_B), \boldsymbol{B})\}_{K_A}$ is underivable from message 6. The proposed fix using message $6'$ adds the sender ID (B), as part of the encrypted message. Thus,

the statement $\{\mathcal{S}((N_A, N_B), B)\}_{K_A}$ is derivable. This example highlights the value of integrating message-sender construct into Message-meaning Rule.

8.4.2 Aziz-Diffie Protocol

Now, let us analyze the protocol by Aziz and Diffie [14] which was still broken despite the use of the original BAN Logic by the authors to verify it.

The protocol uses public-key cryptography for securing the wireless link between a mobile (M) and a base (B). In the following, *alg_list* denotes a list of flags representing potential secret-key algorithms chosen by M. The flag *sel_alg* represents the particular algorithm selected by B from the list *alg_list*, and is to be employed to encrypt the subsequent data call. The protocol for providing the connection setup between M and B is as follows ([24]):

1. $M \rightarrow B:\ Cert(M),\ N_M,\ alg\_list$
2. $B \rightarrow M:\ Cert(B),\ \{X_B\}_{K_M},\ sel\_alg,\ \{hash(\{X_B\}_{K_M}, sel\_alg, N_M, alg\_list)\}_{K_B^{-1}}$
3. $M \rightarrow B:\ \{X_M\}_{K_B},\ \{hash(\{X_M\}_{K_B}, \{X_B\}_{K_M}\}_{K_M^{-1}}$

Here N_M is a nonce from M, whereas $Cert(M)$ and $Cert(B)$ are certificates of M and B respectively. X_B and X_M denote the particular session key values chosen by B and M, respectively. The final session key x is calculated as $X_M \oplus X_B$.

The protocol was verified in [14] using BAN Logic. Our analysis however reveals that the given proof in [14] apparently contains a flaw. The flaw is introduced in the error-prone idealization step of the formalism. The work [14] idealized message 2 (using the original BAN Logic) as:

$$\{\{\overset{K_B}{\longmapsto} B\}_{K_{Ca}^{-1}}, \boldsymbol{M} \overset{\boldsymbol{X_B}}{\longleftrightarrow} \boldsymbol{B}, N_M\}_{K_B^{-1}}. \tag{8.28}$$

The problem with this is that what can actually be derived from the message above is: $M \overset{\{X_B\}_{K_M}}{\longleftrightarrow} B$ or $\{M \overset{X_B}{\longleftrightarrow} B\}_{K_M}$, and not $M \overset{X_B}{\longleftrightarrow} B$. This formalism pitfall allowed [14] to incorrectly derive the desired goals despite the loophole in the protocol.

After the publication of [14], both [98] and [24] managed to mount an attack on the protocol. The attack outlined in [24] makes use of two parallel open sessions. Impersonating M in the first session, an attacker C is able to obtain $\{X_B\}_{K_M}$ from message 2. In the second session, C then replays $\{X_B\}_{K_M}$ to the initiating M when it plays a role as B. The work [24] correctly pointed out the source of the problem is that C can construct message 2 without the knowledge of X_B. To fix the protocol, [24] proposed the modifications of message 2 and 3 into:

$2'.\ B \rightarrow M:\ Cert(B),\ \boldsymbol{N}_B,\ \{X_B\}_{K_M},\ sel\_alg,\ \{hash(\boldsymbol{X}_B, \boldsymbol{M}, N_M, alg\_list)\}K_B^{-1}$
$3'.\ M \rightarrow B:\ \{X_M\}_{K_B},\ \{hash(\boldsymbol{X}_M, \boldsymbol{B}, \boldsymbol{N}_B\}K_M^{-1}$

Nonce $\boldsymbol{N}_B$ now provides the freshness assurance, taking the role of $\{X_B\}_{K_M}$ in the original protocol. Note that message $2'$ contains $\boldsymbol{M}$ in the signed hash's arguments. Likewise, message $3'$ now contains $\boldsymbol{B}$. Such inclusions are thus in line with our requirement of *message-recipient* in Message-meaning (for signed message) Rule. Although M and B are not included in the clear portions of message $2'$ and $3'$ respectively, they are actually present from the message transfer context. As such, we should take their presence into account in the idealized protocol.

In message $2'$, the encrypted message portion $\{X_B\}_{K_M}$ is sent with a sender assurance outside of the encrypted message portion. The assurance is implicitly established by the sender's knowledge of X_B in the message's signature. Thus, this is the case of sender-assurance outside of the encrypted message ($\mathcal{S}(\{X_B\}_{K_M}, B)$) analyzed in Section 8.3.10. Therefore, we can process the statement using the Message-meaning for encrypted message portion (with external stated-sender information) Rule defined in (8.20). Similarly, message $3'$ employs the same technique on $\{X_M\}_{K_B}$. An alternative proposed solution is to use (explicit) enclosed sender-stated assurances. In this case, we can have $\{X_B, B\}_{K_M}$ in message $2'$, and $\{X_M, M\}_{K_B}$ in message $3'$.

8.5 Sample Application of MPKI-BAN Logic

Appendix E shows how we can apply MPKI-BAN Logic to formally analyze a PKI-based protocol. We have chosen the *3-way Session-Establishment* protocol of CREV-II outlined in Section 7.3.2. The protocol is chosen here so that we can establish the security assurance of the proposed session establishment of hash-chaining based revocation service. The idealized protocol, protocol goals, assumptions, and the proof that the protocol can achieve the goals are given in Appendix E.

8.6 Discussion

We have presented MPKI-BAN as an extension of BAN Logic which deals with PKI. It solves a number of issues in GS-BAN [48] whose solutions are vital to a more accurate reasoning with certificate-based protocols. In summary, we have made the following main contributions.

- We present an improved idealization of certificate, in which the assurance of a private key is also derived from certificate. This eliminates the need to have an assumption about the goodness of the other principal's private key as in GS-BAN.

- We define a new Message-meaning for (private-key) signed message Rule, which contains the message-recipient construct. By doing so, we manage to eliminate the

requirement for stating a message-recipient as a goal, as well as that for introducing an assumption about the sender's jurisdiction over message-recipient as in GS-BAN.

- We also modify the Certificate-validation Rule, which now includes a third premise to highlight the need for a certificate revalidation step. In the CRL model, the new rule thus makes explicit of two requirements in certificate validation: time synchronization with the certificate issuer and checking with the issuer's recent CRL.

- We define the Message-meaning for (public-key) encrypted message Rule which now requires a message-sender construct. This modification is vital as it prevents a common pitfall in public-key protocol design, as clearly illustrated among others by Lowe's attack on Needham-Schroeder public-key protocol [89].

Although some of the modifications may look simple, they are however crucial for better reasoning with PKI-based protocols. Table 8.1 contrasts several constructs and rules from GS-BAN with those in our new MPKI-BAN.

8.7 Conclusion

We have presented MPKI-BAN, which makes BAN Logic more in line with concepts and practices in the modern PKI setting. MPKI-BAN removes various limitations of [48], which we refer to as GS-BAN, for more concise reasoning with a certificate-based public-key authentication. In particular, MPKI-BAN redresses the reasoning on the goodness of private keys, and takes certificate revocation into account. Furthermore, it also addresses common pitfalls in public-key based protocol design due to insufficient attention placed on the "intended recipient" and "stated sender" of a message. MPKI-BAN makes the recipient and the sender explicit, and these requirements are tightly coupled into the logic. This way, it reduces the likelihood of allowing such flaws in the protocol and the corresponding formalism. In addition, our logic also deals with the cases of (public-key) signed encrypted message and encrypted signed message. Some examples on the usage of MPKI-BAN are given, and these examples demonstrate that MPKI-BAN can help prevent common mistakes in public-key protocol design and verification.

Given that BAN Logic is a widely-used logic, we envisage that MPKI-BAN extension can provide a practical and valuable tool without requiring the users to manipulate a substantially complex formalism. It is indeed our aim to make the formal analysis on PKI-based protocols become more handily accessible to a wider range of protocol designers, thus allowing more parties to improve the security of their protocols.

Construct/Rule	**GS-BAN Extension [48]**	**Our MPKI-BAN Extension**
Idealized certificate	$\sigma((\Theta(t_1^P, t_2^P), \wp\kappa\,(P, K_P)), K_I^{-1})$	$\sigma(\Re(\Theta'(t_1^P, t_2^P, I, (\wp\kappa\,(P, K_P), \Pi(K_P^{-1}))), all), K_I^{-1})$
Certificate-validation	$\dfrac{P \models Q \mid\sim (\Theta(t_1^R, t_2^R), C^R), P \models Q \models \Delta(t_1^R, t_2^R)}{P \models Q \models C^R}$	$\dfrac{P \models Q \mid\sim \Theta'(t_1^R, t_2^R, Q, C^R), P \models Q \models \Delta(t_1^R, t_2^R), P \models Q \models \Phi(C^R)}{P \models Q \models C^R}$
Message-recipient	$\Re(X, P)$	$\sigma(\Re(X, P), K_Q^{-1})$
Stated-sender	-	$\{\mathcal{S}(X, Q)\}_{K_P}$ or $\mathcal{S}(\{X\}_{K_P}, Q)$
Message-meaning for signed message	$\dfrac{P \models \wp\kappa\,(Q, K_Q)\,, P \models \Pi(K_Q^{-1}), P \triangleleft \sigma(X, K_Q^{-1})}{P \models Q \mid\sim X}$	$\dfrac{P \models \wp\kappa\,(Q, K_Q)\,, P \models \Pi(K_Q^{-1}), P \triangleleft \sigma(\Re(X, P), K_Q^{-1})}{P \models Q \mid\sim X}$
Message-meaning for signed message (after decryption)	-	$\dfrac{P \models \wp\kappa\,(Q, K_Q)\,, P \models \Pi(K_Q^{-1}), P \mid\sim \sigma(\Re(X, P), K_Q^{-1})}{P \models Q \mid\sim X}$
Message-meaning for encrypted message	-	$\dfrac{P \models \wp\kappa\,(P, K_P)\,, P \models \Pi(K_P^{-1}), P \triangleleft \{\mathcal{S}(X, Q)\}_{K_P}}{P \models Q \mid\sim X}$
Message-meaning for encrypted message portion (after Message-meaning for signed message):		
with included stated-sender and included freshness assurance	-	$\dfrac{P \models \wp\kappa\,(P, K_P)\,, P \models \Pi(K_P^{-1}), P \models Q \mid\sim \{\mathcal{S}(X, Q)\}_{K_P}}{P \models Q \mid\sim X}$
with included stated-sender and external freshness assurance	-	$\dfrac{P \models \wp\kappa\,(P, K_P)\,, P \models \Pi(K_P^{-1}), P \models Q \mid\sim \{\mathcal{S}(X, Q)\}_{K_P}, Y}{P \models Q \mid\sim X, Y}$
with external stated-sender and included freshness assurance	-	$\dfrac{P \models \wp\kappa\,(P, K_P)\,, P \models \Pi(K_P^{-1}), P \models Q \mid\sim \mathcal{S}(\{X\}_{K_P}, Q)}{P \models Q \mid\sim X}$
with external stated-sender and external freshness assurance	-	$\dfrac{P \models \wp\kappa\,(P, K_P)\,, P \models \Pi(K_P^{-1}), P \models Q \mid\sim \mathcal{S}(\{X\}_{K_P}, Q), Y}{P \models Q \mid\sim X, Y}$

Table 8.1: Comparison of public-key constructs and rules from GS-BAN [48] and those from our MPKI-BAN.

Chapter 9

Conclusion

This chapter concludes this thesis. We summarize the contributions of this thesis in Section 9.1, and discuss some future directions in Section 9.2.

9.1 Summary of the Thesis

Malware is a critical security threat today, particularly to connected host systems. It has now evolved from occasional exploits to a global multi-million dollar criminal industry, and is threatening the Internet economy. On the other hand, computer systems are much more susceptible to attacks now than in the past, due to the increasing connectivity, complexity and extensibility.

In this thesis, we focus on protecting software, i.e. programs, to ensure that they run as intended without violating a host's security. We safeguard software from various attack vectors based on the Program Protection Life Cycle (PPLC) framework (see Figure 1.1). Our works reported in this thesis aim to mitigate malware threats by deploying additional security mechanisms on top of the OS as well as securing critical infrastructure which supports the host's security mechanisms. We have shown how we can enhance host security using the proposed security measures along the PPLC stages. These measures provide effective defense against numerous attack vectors while incurring acceptable performance overheads. Here, we summarize the main conclusions from our results.

Firstly, to protect a running program on a host particularly against zero-day attacks, we have conducted an in-depth security analysis on the Self-based (Stide) IDS model [60, 140, 141] and related system-call monitoring IDSs such as [136]. We have presented an efficient algorithm for automated mimicry attack construction on the Self-based (Stide) IDS and its more precise graph-based variant. Using several real programs and exploits, we also have shown the practicality of the attack construction, with execution times of at most a few seconds. We also have shown how the construction method can be generalized into an approach to evaluate the robustness of an IDS against attacks targeting the IDS.

The result allows us to find out how computationally expensive it is to perform a search to craft attacks on the IDS. This result shows the feasibility of obtaining quantitative measurements on the IDS resistance against targeted attacks.

Due to threat from mimicry attacks, we also have proposed an improved IDS model by employing an abstraction technique on process credentials and system call arguments. The proposed technique can be easily combined with other system-call based IDS models. To avoid increasing the false positives, a user-supplied specification is used to abstract the arguments and credentials. This specification takes into account security semantics of the operations and their potential effects to system security. Using sample programs, we have shown that the robustness of the IDS is increased when the abstraction is used. We also have some evidences that the increase in detection accuracy does not lead to more false positives. Unlike other data-flow IDS models [83, 112, 155, 22], our IDS enhancement highlights the virtue of incorporating a host's *security model* into the IDS model to result in a more concise IDS.

By monitoring parameterized system calls of a running process, our IDS provides an effective approximation mechanism to ensure *unaltered execution* of a program P. Such monitoring can be done in a fast and online fashion, and it also allows for a possible prevention of suspected system-call invocations(s). Given a belief that P is good for execution as established by Steps 1–3 of the PPLC, and a belief on P's unaltered execution provided by the IDS, a host can establish the desired belief that program P *runs as intended*. That is, a host H believes that a program P performs the operations as intended by its (trusted) developer D, and that the operations will not violate H's security policies. Using BAN Logic notation, we can describe the belief inference by a host H as follows:[1]

$$\frac{H \models \text{good\_for\_execution\_of\_P},\ H \models \text{unaltered\_execution\_of\_P}}{H \models \text{as\_intended\_execution\_of\_P}}. \tag{9.1}$$

Secondly, to secure program invocations on a host, we have demonstrated the practicality of a mandatory, in-kernel, pre-execution binary (executable) authentication mechanism in Windows. We have presented a prototype, BinAuth, which exemplifies a lightweight binary integrity scheme and its integration into an OS. Using BinAuth, we have shown how a mandatory in-kernel authentication system can provide an increased trust on good software execution, while incurring acceptable performance overheads even on a complex OS (Windows). We have also compared BinAuth with a variety of authentication systems [74, 8, 170, 162, 28, 18, 55, 132, 174] using our developed analysis framework, which inspects key design factors for a binary integrity system.

Given that successful malware attacks usually perform modification/addition into a

[1]Symbol '$\models$' denotes *believes* operator (see Section 2.6)

victim host's file system, the proposed practical binary authentication system is a crucial enabler for providing more secure program execution environments on commodity OSes. Using BAN Logic notation, we can state the derivation of H's belief on the integrity of a program P as ensured by BinAuth as follows:

$$\frac{H \models \text{intact\_content\_of\_P},\ H \models \text{intact\_pathname\_of\_P}}{H \models \text{intact\_P}}. \tag{9.2}$$

Thirdly, we have proposed a framework for a machine-oriented vulnerability database, which allows for an automated vulnerability checking on a host using published advisories. We have demonstrated a proof-of-concept database, which shows effective integration of data from multiple sources and can be used directly by a vulnerability scanner. A corresponding scanner prototype has also been developed for Unix/Linux systems, although the basic principles are applicable to other OSes.

Our work was developed and published in [148] when vulnerability representation and automated processing were still not widely addressed by the security community. Given the present on-going standardization efforts on vulnerability processing by the U.S. Government sponsored organizations, such as CVE [105], OVAL [106], CPE [104] and SCAP [115], we are pleased to see that our proposed mechanisms are in line with these efforts. The developed system for automated vulnerability processing on a host is important to prevent any known-vulnerability attacks during a program's unpatched time interval. It helps the host achieve the following important belief:

$$H \models \text{no\_known\_vulnerability\_on\_P}\,. \tag{9.3}$$

Next, to provide an efficient and secure PKI-based program distribution and interaction, we have addressed the need for a lightweight and practical near real-time certificate revocation scheme. We have proposed a family of lightweight, practical and inherently-distributed certificate revocation schemes, called CREV, which are based on the recently available Extended-Validation (EV) certificate infrastructure. Our CREV schemes enhance CRS/NOVOMODO [100, 101] and OCSP [113] to support efficient revocation with near real-time timeliness guarantees (1–10 minutes). We also have developed a realistic model and cost analysis framework for evaluating the certificate revocation schemes. We enhances the model in [91, 65] which make use of the empirical revocation data. Our framework incorporates a number of novel aspects to result in a more accurate model over the original one in [91, 65]. Among others, our model can properly deal with fine-grained certificate generation and CRL issuance. Using the developed model, we have demonstrated that the proposed CREV schemes keep the overheads manageable on all the involved entities, while maintaining lightweight requirements on the verifier even under a timeliness guarantee of 1 minute. The scheme has an additional advantage in that status queries are protected from external parties including the CA, which may be of

importance to the EVCP and the verifiers due to privacy reasons.

The proposed revocation schemes will provide a host with an efficient way to validate PKI-based digital signatures. With respect to PPLC, the host will be able to better ensure a secure program distribution, which is the first step of PPLC. More specifically, the host can ensure that program P actually originates from trusted software developer D, P is received unmodified, and P is a good behavior program.[2] By "*good behavior*", we mean that the program is non-malicious in nature, and has no intention to violate any security policies of the target hosts (beyond the program's known functionalities) or any acceptable use policies. In the language of BAN Logic, the PKI and efficient revocation service enable a host H to derive:

$$H \models P \quad \text{and} \quad H \models \text{good\_behavior\_P}\,. \tag{9.4}$$

Combined with a belief established by the binary authentication system (9.2) and that from the vulnerability-free execution (9.3), we thus can infer the following:

$$\frac{H \models P,\ H \models \text{good\_behavior\_P},\ H \models \text{intact\_P}, H \models \text{no\_known\_vulnerability\_on\_P}}{H \models \text{good\_for\_execution\_of\_P}} \tag{9.5}$$

Finally, we have extended BAN Logic [25, 26] and a prior extension work by Gaarder and Snekkenes [48] in order to make BAN Logic more concise in reasoning on PKI-based protocols. In particular, our extension redresses the reasoning on the goodness of private keys, and considers certificate revocation as a part of certificate verification. We also address common pitfalls in public-key based protocol design due to insufficient attention placed on the "intended recipient" and the "stated sender" of a message. Our extension makes the recipient and the sender explicit, and these requirements are incorporated into the logic. In this way, our logic reduces the likelihood of allowing such vulnerabilities in the protocol and the corresponding formalism. In addition, our logic also deals with cases of (public-key) signed encrypted messages and encrypted signed messages. We also have shown the usage and the sample application of our logic to PKI-based protocols, including a session establishment protocol proposed in CREV-I scheme.

Altogether, we have shown how we can safeguard the four PPLC stages with efficient security measures. These proposed measures substantially enhance host security while only incurring acceptable performance overheads. These measures helps the host establish the desired belief: "$H \models$ as_intended_execution_of_P" as shown in (9.1). In summary, our proposed techniques help provide strong additional layers of protection to securing program execution. This can help address timely and pressing problem of the increasing malware threats, particularly to today's connected host systems.

[2] In addition to its content, P also carries with itself a belief statement that it is a good behavior program. This belief is accepted by host H when it decides to install P into its system at time $t_{installed}$.

9.2 Future Work

We have identified the following directions to be pursued in the future.

Resource access monitoring on a program

One weakness of the PPLC model is that a host needs to trust a software developer that a piece of software behaves in a good manner and will not violate the host's security policies. One way to deal with this limitation is by adding *resource access monitoring of a program* as an additional step to PPLC. The host can generate the resource-access profile of a program by observing the program's normal usage. Alternatively, the profile can be supplied by the developer based on the program's behavior definition in terms of resource access patterns. In this way, a host H can limit (sandbox) a program execution within a set of approved operations. As a result, host H can believe that not only a program P runs as intended (i.e. "$H \models$ as_intended_execution_of_P"), but also within a specified capabilities in terms of the host's resources accessed (i.e. "$H \models$ within_access_specification_execution_of_P").

Richer IDS model and application to other OSes

For an anomaly detector IDS, we can combine the gray-box IDS model developed from observing system call invocations with the white-box techniques if the program's source code is available [87]. Alternatively, binary analysis techniques [142] can be used to derive a more concise IDS model. Binary instrumentation technique can also be used to "randomize" normal system-call profile generations on a host. This would render the attacker's approximated normal profile different from the valid (randomized) one and thus less effective. With regard to our IDS attack-construction framework, we also intend to apply it to more IDS models.

We also would like to implement the PAC-based IDS for other OSes, such as Windows. Additionally, the PAC-based IDS model can also be tailored more for specific applications, such as web applications [57, 37].

More usable and robust binary authentication system

We can extend the BinAuth system to protect not only binary files, but also important non-binary files. Additionally, it can be extended to deal with various usage scenarios that require special treatment of binary authentication, such as during software development, software installations and software updates.

BinAuth does prevent untrusted binaries from being executed. It however does not prevent illegal additions or modifications to protected binaries. One possible work is to make BinAuth also monitor illegal modifications on protected binaries. Such monitoring, however, needs to be justified with respect to the incurred overheads. Given the availability of a secure booting infrastructure, such as Trusted Platform Module (TPM), we

can also try to combine BinAuth with TPM. This way, we can better ensure the integrity of the overall protection chain, starting from the hardware to the applications. Some keys and important files used by BinAuth can also be securely stored by TPM.

Richer vulnerability description model and improved implementation

To ensure a vulnerability-free host system, we would like to enhance the Movtraq framework by devising a richer vulnerability description model. The model should capture a vulnerability in a more precise way, yet still be amenable for automated processing. Another direction worth exploring is to standardize vulnerability checks on a target machine for different popular OSes. Such checks could also be parameterized to make them as generic as possible. We would like to make these checks represented as database entries, thus allowing extensible vulnerability checks definitions. An entry could specify the description of how the check should be performed on different OSes, database fields for the values of objects involved in the checking, and symbolic description of the pre-conditions and consequence as in Movtraq.

Our Movtraq vulnerability database uses SQL for data access and transfer. To be in line with recent standardization efforts such as OVAL, we can redesign the database format to use an XML-based data representation. One of the limitations of OVAL framework is its lack of specifications for vulnerability pre-requisites and consequences to allow vulnerability chain reasoning [92, 33]. We thus can extend OVAL by specifying a vulnerability's pre-requisites and consequences using a symbolic-description approach taken in Movtraq. Other possible enhancements on Movtraq implementation could include convenient GUIs, non Perl-based scanner, and compatibility with various popular OSes.

Analysis of certificate revocation schemes

In the analysis of certificate revocation schemes using our realistic framework, we can calculate the overheads of various other revocation schemes to be compared with the CREV schemes. We can further improve the developed framework by allowing certificates to have different lifetimes. Lastly, we can also conduct a simulation to support the analysis results derived using the framework.

Formal analysis on PKI-based protocols

In proposing MPKI-BAN Logic, we take an approach which focuses on pragmatism. We center on the logic definition and its usage application while leaving theoretical analysis of the logic, such as the logic's soundness and completeness, as a separate treatment beyond our work's scope. Developing a semantic model for MPKI-BAN is an important future work. Additionally, we can try to apply public-key constructs and rules developed in MPKI-BAN to other authentication logics, as well as non modal-logic based formal analysis techniques.

Appendix A

Sample Configuration for Privilege and Argument Categorization

```
#====== Sample Configuration File =======#
# EUID Abstraction Section
# Format:  <categorized-euid>:<euid1>,<euid2>,...
0:0
1:2000,2001,2003
2:3000
100:*
----------------------------------------
# EGID Abstraction Section
# Format:  <categorized-egid>:<egid1>,<egid2>,...
0:0
1:1,2,3,5
100:*
----------------------------------------
# Argument Abstraction Section
# Format:  <syscall> <arg1> <arg2> <arg3> ...  <cat-value>
open p=/etc/passwd o=O_WRONLY|o=O_RDRW * 1
open p=/etc/shadow o=O_WRONLY|o=O_RDRW * 2
open p=/etc/group o=O_WRONLY|o=O_RDRW * 3
open p=/proc/kmem o=O_WRONLY|o=O_RDRW * 4
open p=/etc/hosts.equiv o=O_WRONLY|o=O_RDRW * 5
open p=/etc/* o=O_WRONLY|o=O_RDRW * 6
open p=/etc/* o=O_RDONLY * 7
open p=/bin/* o=O_WRONLY|o=O_RDRW * 8
open p=/sbin/* o=O_WRONLY|o=O_RDRW * 9
open p=/boot/* o=O_WRONLY|o=O_RDRW * 10
open p=/dev/* o=O_WRONLY|o=O_RDRW * 11
open p=/usr/* o=O_WRONLY|o=O_RDRW * 12
open p=/lib/* o=O_WRONLY|o=O_RDRW * 13
open p=/var/* o=O_WRONLY|o=O_RDRW * 14
open p=/mnt/* o=O_WRONLY|o=O_RDRW * 15
open p=/proc/* o=O_WRONLY|o=O_RDRW * 16
open * o=O_WRONLY|o=O_RDRW * 17
open * * * 18
```

```
chmod p=/etc/{passwd,shadow,group,hosts.equiv}|p=/proc/kmem * 1
chmod * * - 2
fchmod p=/etc/{passwd,shadow,group,hosts.equiv}|p=/proc/kmem * 1
fchmod * * - 2
chown p=/etc/{passwd,shadow,group,hosts.equiv}|p=/proc/kmem * 1
chown * * * 2
fchown p=/etc/{passwd,shadow,group,hosts.equiv}|p=/proc/kmem * 1
fchown * * * 2
lchown p=/etc/{passwd,shadow,group,hosts.equiv}|p=/proc/kmem * 1
lchown * * * 2
rename p=/etc/{passwd,shadow,group} * 1
rename p=/proc/kmem * 2
rename p=/etc/hosts.equiv * 3
rename * p=/etc/{passwd,shadow,group} 4
rename * p=/proc/kmem 5
rename * p=/etc/hosts.equiv 6
rename * * 7
link p=/etc/{passwd,shadow,group} * 1
link p=/proc/kmem * 2
link p=/etc/hosts.equiv * 3
link * p=/etc/{passwd,shadow,group} 4
link * p=/proc/kmem 5
link * p=/etc/hosts.equiv 6
link * * 7
unlink p=/etc/{passwd,shadow,group,hosts.equiv}|p=/proc/kmem 1
unlink * 2
symlink p=/etc/{passwd,shadow,group} * 1
symlink p=/proc/kmem * 2
symlink p=/etc/hosts.equiv * 3
symlink * p=/etc/{passwd,shadow,group} 4
symlink * p=/proc/kmem 5
symlink * p=/etc/hosts.equiv 6
symlink * * 7
mount * p=/etc * * * 1
mount * p=/{bin,sbin} * * * 2
mount * p=/boot * * * 3
mount * p=/usr * * * 4
mount * p=/lib * * * 5
mount * p=/proc * * * 6
mount * * * * * 7
mknod p=/etc/{passwd,shadow,group,hosts.equiv}|p=/proc/kmem * * 1
mknod * * * 2
init_module * * * 1
execve p=/bin/{sh,csh,ksh} * * 1
execve * * * 2
----------------------------------------
# Illegal Transitions Section
# Format:  <syscall> <cat-values> [<cat-euid>,<cat-egid>]*
open [1..6,8-11,13-15] 0,* *,0
{chmod,fchmod,chown,fchown,lchown,mknod,unlink,init_module,execve} 1 0,* *,0
{rename,link,symlink,mount} [1..6] 0,* *,0
----------------------------------------
```

Appendix B

Database Entities in Movtraq Vulnerability Database

Movtraq vulnerability database (see Section 6.3) defines seven entities (tables) to store the information of a vulnerability entry. Their key fields, as seen from an integration and machine processing perspective, are listed below.

Vulnerability Entity — records main information about each vulnerability: Vul_ID, Vulnerability name, CVE_ID, other corresponding IDs (e.g. CERT_ID, BugTraq_ID), a textual description for the vulnerability, and other fields as keys relating to other tables (i.e. Vulnerability Specifications and Environment Specifications Entity).

Vulnerability Specifications Entity — records component factors: vulnerability consequences*, key relating to Operating System, Application, and Services Entity.

Operating System Entity — records component factors originating from the OS: OS ID, OS name, vulnerable versions†, hardware type information†.

Application Entity — records component factors due to a vulnerable application: application ID, name of application, vulnerable versions†, hardware type information†.

Services Entity — records component factors due to services (a service could be a daemon): service ID, name of service, vulnerable versions†, hardware type information†, protocols, and port numbers.

Environment Specifications Entity — records environmental factors: existence of required user/application/service object*, existence of a file object*, kernel versions of the running OS†, remote exploitation flag.

Exploit Entity — records details of any available exploits: exploit information (URL, filename, description, etc.), privileges needed for the exploit*, consequences of using the exploit*.

The fields labeled by '*' make use of the vulnerability description expressions or vulnerability target objects described in Section 6.4. Those labeled by '†' can use the $Software\_ID$ (see Section 5.6) or the CPE [104] format. Note that some fields sharing a similar function can occur a few times in different contexts. *Consequence* field, for example, occurs both in Vulnerability Specifications Entity and Exploit Entity. The former is the consequence of the specified vulnerability, whereas the latter is from using a specific available exploit.

Appendix C

Relevant Rules of BAN Logic

Below are the rules of the original BAN Logic [25, 26] which are relevant to our MPKI-BAN extension explained in Chapter 8. We follow the notation and description of BAN Logic as in [26].

- (R_1) **Message-meaning (for secret-key encryption) Rule**:
  If P believes that the key K_{PQ} is shared with Q, and sees a message X encrypted under K_{PQ}, then P believes that Q once said X.)

$$\frac{P \models Q \overset{K_{PQ}}{\longleftrightarrow} P, P \triangleleft \{X\}_{K_{PQ}}}{P \models Q \mid\sim X}$$

- (R_2) **Nonce verification Rule**:
  (This checks that a message is recent, and hence that the sender still believes in it.)

$$\frac{P \models \sharp(X), P \models Q \mid\sim X}{P \models Q \models X}$$

- (R_3) **Jurisdiction Rule**:
  (If P believes that Q has jurisdiction over X, then P trusts Q on the truth of X.)

$$\frac{P \models Q \Rightarrow X, P \models Q \models X}{P \models X}$$

- (R_4) **And-Introduction Rule**:
  (If P believes each individual statement separately, then P believes a set of statements.)

$$\frac{P \models X, P \models Y}{P \models (X, Y)}$$

- (R_5) **And-Elimination Rule**:
  (If P believes a set of statements, then P believes each individual statement separately.)

$$\frac{P \models (X, Y)}{P \models (X)}$$

- (R_6) **Believe And-Elimination Rule**:
(And-elimination rule that applies to a belief held by other principal.)
$$\frac{P \models Q \models (X, Y)}{P \models Q \models X}$$

- (R_7) **Said And-Elimination Rule**:
(And-elimination rule that applies to the $\mid\sim$ operator. Note that and-introduction rule does not apply to $\mid\sim$.)
$$\frac{P \models Q \mid\sim (X, Y)}{P \models Q \mid\sim X}$$

- (R_8) **See components Rule**:
(If P sees a formula, then P also sees its components, provided he knows the necessary keys.)
$$\frac{P \triangleleft (X, Y)}{P \triangleleft X}$$

- (R_9) **See encrypted message (for secret-key) Rule**:
(If P sees an message encrypted with a secret key K_{PQ}, and that he knows the key K_{PQ}, then P also sees its message.)
$$\frac{P \models Q \overset{K_{PQ}}{\longleftrightarrow} P, P \triangleleft \{X\}_{K_{PQ}}}{P \triangleleft X}$$

- (R_{10}) **Freshness extension Rule**:
(If one part of a formula is known to be fresh, then the entire formula is also fresh.)
$$\frac{P \models \sharp(X)}{P \models \sharp(X, Y)}$$

- (R_{11}) **Key-symmetry (for secret-key) Rule**:
(A secret key is used between a pair of principals in either direction.)
$$\frac{P \models R \overset{K_{RR'}}{\longleftrightarrow} R'}{P \models R' \overset{K_{R'R}}{\longleftrightarrow} R}$$

- (R_{12}) **Believe Key-symmetry (for secret-key) Rule**:
(Key-symmetry that applies to a belief held by other principal.)
$$\frac{P \models Q \models R \overset{K_{RR'}}{\longleftrightarrow} R'}{P \models Q \models R' \overset{K_{R'R}}{\longleftrightarrow} R}$$

Appendix D

New Rules of MPKI-BAN Logic

Below are the rules newly defined in our MPKI-BAN Extension Logic in Chapter 8.

- (R_{13}) **New Message-Meaning for (private-key) signed message Rule**:

$$\frac{P \models \wp\kappa\left(Q, K_Q\right), P \models \Pi(K_Q^{-1}), P \triangleleft \sigma(\Re(X, P), K_Q^{-1})}{P \models Q \mid\sim X}$$

- (R_{14}) **All-Recipient See Rule**:

$$\frac{P \triangleleft \sigma(\Re(X, all), K_Q^{-1})}{P \triangleleft \sigma(\Re(X, P), K_Q^{-1})}$$

- (R_{15}) **New Certificate-Validation Rule**:

$$\frac{P \models Q \mid\sim \Theta'(t_1^R, t_2^R, Q, C^R), P \models Q \models \Delta(t_1^R, t_2^R), P \models Q \models \Phi(C^R)}{P \models Q \models C^R}$$

  with $C^R = \wp\kappa\left(P, K_P\right), \Pi(K_P^{-1})$.

- (R_{16}) **Duration-Stamp (without revocation) Validation Rule**:

$$\frac{P \models Q \mid\sim \Theta(t_1, t_2, X), P \models Q \models \Delta(t_1, t_2)}{P \models Q \models X}$$

- (R_{17}) **New Message-Meaning for (public-key) encrypted message Rule**:

$$\frac{P \models \wp\kappa\left(P, K_P\right), P \models \Pi(K_P^{-1}), P \triangleleft \{\mathcal{S}(X, Q)\}_{K_P}}{P \models Q \mid\sim X}$$

- (R_{18}) **Message-Meaning for (private-key) signed message – after decryption) Rule**:

$$\frac{P \models \wp\kappa\left(Q, K_Q\right), P \models \Pi(K_Q^{-1}), P \mid\sim \sigma(\Re(X, P), K_Q^{-1})}{P \models Q \mid\sim X}$$

- (R_{19}) **Message-Meaning for encrypted message portion – after Message-Meaning for signed message, with included stated-sender and included freshness assurance Rule**:

$$\frac{P \models \wp\kappa\left(P, K_P\right), P \models \Pi(K_P^{-1}), P \models Q \mid\sim \{\mathcal{S}(X, Q)\}_{K_P}}{P \models Q \mid\sim X}$$

- (R_{20}) **Message-Meaning for encrypted message portion – after Message-Meaning for signed message, with included stated-sender and external freshness assurance Rule**:

$$\frac{P \models \wp\kappa\left(P, K_P\right), P \models \Pi(K_P^{-1}), P \models Q \mid\sim \{\mathcal{S}(X, Q)\}_{K_P}, Y}{P \models Q \mid\sim X, Y}$$

- (R_{21}) **Message-Meaning for encrypted message portion – after Message-Meaning for signed message, with external stated-sender and included freshness assurance Rule**:

$$\frac{P \models \wp\kappa\left(P, K_P\right), P \models \Pi(K_P^{-1}), P \models Q \mid\sim \mathcal{S}(\{X\}_{K_P}, Q)}{P \models Q \mid\sim X}$$

- (R_{22}) **Message-Meaning for encrypted message portion – after Message-Meaning for signed message, with external stated-sender and external freshness assurance Rule**:

$$\frac{P \models \wp\kappa\left(P, K_P\right), P \models \Pi(K_P^{-1}), P \models Q \mid\sim \mathcal{S}(\{X\}_{K_P}, Q), Y}{P \models Q \mid\sim X, Y}$$

- (R_{23}) **Message-Meaning (for secret-key hashed message) Rule**:

$$\frac{P \models Q \overset{K_{PQ}}{\longleftrightarrow} P, P \triangleleft \mu(X, K_{PQ})}{P \models Q \mid\sim X}$$

- (R_{24}) **See Hashed-Message Rule**:

$$\frac{P \triangleleft \mu(X, K_{PQ})}{P \triangleleft X}$$

- (R_{25}) **See Signed-Message Rule**:

$$\frac{P \triangleleft \sigma(X, K_Q^{-1})}{P \triangleleft X}$$

- (R_{26}) **See Recipient-tagged Message Rule**:

$$\frac{P \triangleleft \Re(X, P)}{P \triangleleft X}$$

Appendix E

Sample Application of MPKI-BAN Logic

This appendix shows how we can apply MPKI-BAN Logic (elaborated in Chapter 8) to formally analyze a PKI-based protocol. We have chosen the *3-way Session-Establishment* protocol of CREV-I outlined in Section 7.3.2. The protocol is chosen here so that we can establish the security assurance of the proposed session establishment of a hash-chaining based revocation service.

For ease of reference, we show the protocol again here (only the message transfers are shown):

1. $EVCP \rightarrow CA$: "*Hash-Chain Session Establishment Request*"=
   $\langle EVCP\_ID, CA\_ID, Serial\_No, T, nonce_{EVCP}, SigAlgID, Signature\rangle$
   where:
   T = either a timestamp or CA's nonce obtainable from URI_{CA_nonce};
   $SigAlgID$ = identification for the signing algorithm.

2. $CA \rightarrow EVCP$: "*Hash-Chain Session Establishment Reply*"=
   $\langle EstablishmentStatus, CA\_ID, EVCP\_ID, nonce_{EVCP}, Serial\_No, CertStatus,$
   $HashAlgID, d, t_{CREV\_II}, Y, N, SessionStart, SessionExpiry, SigAlgID, Signature\rangle$
   where:
   $HashAlgID$, d, Y, N = hash-chain parameters;
   $SessionStart$ and $SessionExpiry$ = the start and end time of the established session;
   t_{CREV_II}=*SessionExpiry-SessionStart*= session's lifetime.

3. $EVCP \rightarrow CA$: "*Hash-Chain Session Establishment ACK*"=
   $\langle EVCP\_ID, CA\_ID, Serial\_No, SessionStart, SigAlgID, Signature\rangle$.

E.1 Idealized Protocol

We can idealize the messages in the protocol above as follows.

1. $EVCP \rightarrow CA$: $\sigma(\Re\ ((Serial\_No, T, nonce_{EVCP}), CA), K^{-1}_{EVCP})$

2. $CA \rightarrow EVCP$: $\sigma\big(\Re\left((nonce_{EVCP}, Hash\_Chain), EVCP\right), K_{CA}^{-1}\big)$
with: $Hash\_Chain = (EstablishmentStatus,\ Serial\_No,\ CertStatus,\ SessionStart,$
$SessionExpiry,\ HashAlgID,\ d,\ t_{CREV\_II},\ Y,\ N)$

3. $EVCP \rightarrow CA$: $\sigma(\Re\ ((Serial\_No, SessionStart, EVCP \models Hash\_Chain), CA)\ , K_{EVCP}^{-1})$

The message-recipient construct ($\Re(X, receiver)$) is idealized in all messages above. This is because each message always contains the receiver ID, which is also calculated in the signature generation.

We specifically define $Hash\_Chain$ as the message portion containing all the parameters specifying a hash-chaining based (e.g. CRS/NOVOMODO) revocation system. $Hash\_Chain$ is to be used by the verifiers as reference token information, together with the CA's latest hash token, in ensuring the validity of the EVCP's certificate.

Note that the protocol above is between an EVCP and the CA. Thus, unlike a typical protocol where the principals establish a trust on each other through a (common) CA, here it is the CA with whom the EVCP deals directly. Since all principals are assumed to have believed the CA's public and private keys, we do not logically infer the following EVCP's beliefs: $EVCP \models \wp\kappa\left(CA, K_{CA}\right)$ and $EVCP \models \Pi(K_{CA}^{-1})$. Rather, they are assumed to be true at the start of the protocol run. Conversely, since the CA is the authoritative entity on the goodness of the EVCP's public and private key pair, the CA determines by itself the goodness of the keys (i.e. $CA \models \wp\kappa\left(EVCP, K_{EVCP}\right)$ and $CA \models \Pi(K_{EVCP}^{-1})$).

E.2 Initial-State Assumptions

We list all the assumptions below:

- EVCP believes correct and good public and private keys of CA:
  α_1: $EVCP \models \wp\kappa\left(CA, K_{CA}\right)$
  α_2: $EVCP \models \Pi(K_{CA}^{-1})$
- CA believes the goodness of EVCP's public and private keys (based on its certification record):
  α_3: $CA \models \wp\kappa\left(EVCP, K_{EVCP}\right)$
  α_4: $CA \models \Pi(K_{EVCP}^{-1})$
- Each principal believes the freshness of its own nonce or timestamp (or information functioning as a timestamp):
  α_5: $CA \models \sharp(T)$
  α_6: $EVCP \models \sharp(nonce_{EVCP})$
  α_7: $CA \models \sharp(SessionStart)$
- CA believes that EVCP has a jurisdiction over the Session-Establishment Request message portion:
  α_8: $CA \models EVCP \Rightarrow Serial\_No, nonce_{EVCP}$
- EVCP believes that CA has a jurisdiction over $Hash\_Chain$:
  α_9: $EVCP \models CA \Rightarrow Hash\_Chain$

- CA believes that EVCP has a jurisdiction over EVCP's belief on $Hash\_Chain$:
  α_{10}: $CA \models EVCP \Rightarrow (EVCP \models Hash\_Chain)$

E.3 Protocol Goals

The CREV-I session establishment protocol does not aim to establish a session key to secure the subsequent communication session. Rather, it aims to establish a mutual authentication between the CA and EVCP, and to allow EVCP to obtain the CA's $Hash\_Chain$. In addition, the CA needs to establish a belief that EVCP has believed its generated $Hash\_Chain$.

We list the goals as follows:

- CA believes $Serial\_No, nonce_{EVCP}$ (sent in EVCP's *Hash-Chain Session Establishment Request*):
  G_1: $CA \models Serial\_No, nonce_{EVCP}$
- EVCP believes $Hash\_Chain$:
  G_2: $EVCP \models Hash\_Chain$
- CA believes that EVCP believes $Hash\_Chain$:
  G_3: $CA \models (EVC \models Hash\_Chain)$

E.4 The Proof

We prove that the proposed protocol achieve its stated goals as follows. Note that the rule numbering here follows that of Appendix D.

After **message 1**, we have:

$$CA \triangleleft \sigma(\Re((Serial\_No, T, nonce_{EVCP}), CA), K_{EVCP}^{-1}) \tag{E.1}$$

Using Message-meaning for signed message Rule (R_{13}) on α_3, α_4, and (E.1):

$$CA \models EVCP \mid\sim Serial\_No, T, nonce_{EVCP} \tag{E.2}$$

Using Freshness extension Rule (R_{10}) on α_5:

$$CA \models \sharp(Serial\_No, T, nonce_{EVCP}) \tag{E.3}$$

Using Nonce verification Rule (R_2) on (E.3) and (E.2):

$$CA \models EVCP \models Serial\_No, T, nonce_{EVCP} \tag{E.4}$$

Using And-elimination Rule (R_5) on (E.4):

$$CA \models EVCP \models Serial\_No, nonce_{EVCP} \tag{E.5}$$

Using Jurisdiction Rule (R_3) on α_8 and (E.5):

$$\boxed{G_1 : CA \models Serial\_No, nonce_{EVCP}} \tag{E.6}$$

After **message 2**, we have:

$$EVCP \triangleleft \sigma(\Re((nonce_{EVCP}, Hash\_Chain), EVCP), K_{CA}^{-1}) \tag{E.7}$$

Using Message-meaning for signed message Rule (R_{13}) on α_1, α_2, and (E.7):

$$EVCP \models CA \mid\sim nonce_{EVCP}, Hash\_Chain \tag{E.8}$$

Using Freshness extension Rule (R_{10}) on α_6:

$$EVCP \models \sharp(nonce_{EVCP}, Hash\_Chain) \tag{E.9}$$

Using Nonce verification Rule (R_2) on (E.9) and (E.8):

$$EVCP \models CA \models nonce_{EVCP}, Hash\_Chain \tag{E.10}$$

Using And-elimination Rule (R_5) on (E.10):

$$EVCP \models CA \models Hash\_Chain \tag{E.11}$$

Using Jurisdiction Rule (R_3) on α_9 and (E.11):

$$\boxed{G_2 : EVCP \models Hash\_Chain} \tag{E.12}$$

After **message 3**, we have:

$$CA \triangleleft \sigma(\Re\ ((Serial\_No, SessionStart, EVCP \models Hash\_Chain), CA), K^{-1}_{EVCP}) \tag{E.13}$$

Using Message-meaning for signed message Rule (R_{13}) on α_3, α_4, and (E.13):

$$CA \models EVCP \mid\sim (Serial\_No, SessionStart, EVCP \models Hash\_Chain) \tag{E.14}$$

Using Freshness extension Rule (R_{10}) on α_7:

$$CA \models \sharp(Serial\_No, SessionStart, EVCP \models Hash\_Chain) \tag{E.15}$$

Using Nonce verification Rule (R_2) on (E.15) and (E.14):

$$CA \models EVCP \models (Serial\_No, SessionStart, EVCP \models Hash\_Chain) \tag{E.16}$$

Using And-elimination Rule (R_5) on (E.16):

$$CA \models EVCP \models (EVCP \models Hash\_Chain) \tag{E.17}$$

Using Jurisdiction Rule (R_3) on α_{10} and (E.17):

$$\boxed{G_3 : CA \models (EVCP \models Hash\_Chain)} \tag{E.18}$$

E.5 Discussion

We have given a proof of the proposed CREV-I session establishment by using MPKI-BAN. Some interesting points to note from the proof are:

- We can see that the three stated goals (G_1–G_3) are all achievable. Using the protocol, EVCP and the CA can authenticate each other in order to set up a new session. In particular, EVCP can derive a belief on $Hash\_Chain$ (i.e. $G_2 : EVCP \models Hash\_Chain$). Moreover, the CA can also establish a belief that EVCP believes $Hash\_Chain$ (i.e. $G_3 : CA \models (EVCP \models Hash\_Chain)$). Capturing these facts is important because only after establishing these beliefs, the CA then starts sending its periodical hash tokens to EVCP.
- Although the CA believes that EVCP believes $Hash\_Chain$ (i.e. G_3), from the 3-way protocol above, EVCP has no knowledge that the CA has established G_3. In other words, the following belief is not derivable: $EVCP \models CA \models (EVCP \models Hash\_Chain)$. When proposing the protocol, we assume a reliable channel between EVCP and the CA. Hence, EVCP

simply assumes that the CA can derive G_3 upon receipt of message 3 that it has sent. In a lossy channel, mechanism to deal with possible message losses are thus needed.

- We can see that the formalism makes it clearer that message 3 (Hash-Chain Session Establishment ACK) makes use of $SessionStart$ as the freshness assurance. That is, $SessionStart$, generated and sent in message 2 by the CA, functions as a nonce to be returned in message 3 by EVCP. In designing the protocol, we assume that $SessionStart$ is set to the timestamp when the CA approves EVCP's request after verifying message 1. In addition, the CA also keeps track of the latest Hash-Chain Session Establishment ACK message that it sends to each EVCP. Thus, we can ensure that $Serial\_No{+}SessionStart$ is unique. Provided that $SessionStart$ timestamp is sufficiently fine-grained, we can achieve a sufficient protection against a potential oracle attack on the CA's signed message generation. If desired, however, the protocol can be extended to include an additional CA's nonce in message 2. In addition to functioning as a challenge, this nonce also acts as a "salt", which increases the message space so as to reduce the risk of an oracle attack.

Appendix F

List of Author's Published and Submitted Work

The following are published and submitted works by the author during the author's Ph.D. candidature.

Published Works:

- Sufatrio, Roland H. C. Yap and Liming Zhong, "A Machine-Oriented Integrated Vulnerability Database for Automated Vulnerability Detection and Processing", In *Proceedings of the 18th USENIX Large Installation System Administration*, pp. 47–58, 2004.
- Sufatrio and Roland H. C. Yap, "Improving Host-based IDS with Argument Abstraction to Prevent Mimicry Attacks", In *Proceedings of the 8th International Symposium on Recent Advances in Intrusion Detection (RAID)*. pp. 146–164, 2005.
- Rajiv Ramnath, Sufatrio, Roland H. C. Yap, and Wu Yongzheng, "WinResMon: A Tool for Discovering Software Dependencies, Configuration, and Requirements in Microsoft Windows", In *Proceedings of the 20th USENIX Large Installation System Administration*, pp. 175-186, 2006.
- Felix Halim, Rajiv Ramnath, Sufatrio, Yongzheng Wu, and Roland H. C. Yap, "A Lightweight Binary Authentication System for Windows", In *Proceedings of the Joint iTrust and PST Conferences on Privacy, Trust Management and Security (IFIPTM)*, pp. 295–310, Springer, 2008.
- Sufatrio and Roland H. C. Yap, "Extending BAN Logic for Reasoning with Modern PKI-based Protocols", In *Proceedings of the IFIP International Workshop on Network and System Security 2008 (NSS)*, pp. 190–197, 2008.
- Yongzheng Wu, Sufatrio, Roland H. C. Yap, Rajiv Ramnath and Felix Halim, "Establishing Software Integrity Trust: A Survey and Lightweight Authentication System for Windows, Book Chapter, in *Trust Modeling and Management in Digital Environments: From Social Concept to System Development*, Information Science Reference, 2010.

Submitted Work:

- Sufatrio and Roland H. C. Yap, "Practical and Lightweight Certificate Revocation using Extended-Validation Certificate Infrastructure", 2010.

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