On the Semantics of Self-Unpacking Malware Code[*]

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Abstract

The rapid increase in attacks on software systems via malware such as viruses, worms, trojans, etc., has made it imperative to develop effective techniques for detecting and analyzing malware binaries. Such binaries are usually transmitted in packed or encrypted form, with the executable payload decrypted dynamically and then executed. In order to reason formally about their execution behavior, therefore, we need semantic descriptions that can capture this self-modifying aspect of their code. However, current approaches to the semantics of programs usually assume that the program code is immutable, which makes them inapplicable to self-unpacking malware code. This paper takes a step towards addressing this problem by describing a formal semantics for self-modifying code. We use our semantics to show how the execution of self-unpacking code can be divided naturally into a sequence of phases, and uses this to show how the behavior of a program can be characterized statically in terms of a program evolution graph. We discuss several applications of our work, including static unpacking and deobfuscation of encrypted malware and static cross-phase code analysis.

1 Motivation

The rapid increase in the use of the Internet in many aspects of our lives has led to an explosive growth in the spread of malware such as computer worms, viruses, trojans, spyware, and bots. Software security considerations have accordingly become a crucial aspect of modern software design, development, and deployment. Since software—both new applications to be installed on a system and patches or upgrades to existing ones—is commonly distributed in the form of binaries, the ability to verify that a binary file received from elsewhere does not have any malicious content is an important component of software security (for example, email attachments are now routinely subjected to virus scans).

Most malware executables today are transmitted in "scrambled" form. There are two commonly used approaches to such code scrambling: *encryption* and *packing*. The former refers to the use of some kind of invertible operation, together with an encryption key, to conceal the executable code; the latter refers to the use of compression techniques to reduce the size of the malware payload while at the same time converting it to a form where the executable content is hidden. In either case, the scrambled code is "unpacked" at runtime prior to execution; in many cases, there are multiple rounds of unpacking during the course of an execution.[1] Such self-unpacking malware code is, therefore, self-modifying.

In order to build effective defences against malware, we have to be able to obtain a precise understand their behavior. This makes it important to be able to reason formally about the code and the behavior(s) of malware. Such reasoning can be helped greatly by a formal semantics that is able, among other things, to cope with self-modifying code. Unfortunately, current approaches to program semantics typically assume that the program code is immutable, which makes them unsuitable for this purpose. This poses a problem, which this paper attempts to address. We motivate our approach using the example application of static analysis of self-unpacking malware code.

The usual approach to dealing with packed binaries—especially when analyzing malware that has not been

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[1]For most of the paper, we will not be greatly concerned by the technical distinction between the operations of encryption and packing (for changing a program to its scrambled form) on the one hand, and between decryption and unpacking (for the reverse operation) on the other. In order to simplify the presentation, therefore, we will generically use the term *packing* to refer to the former and *unpacking* to refer to the latter.

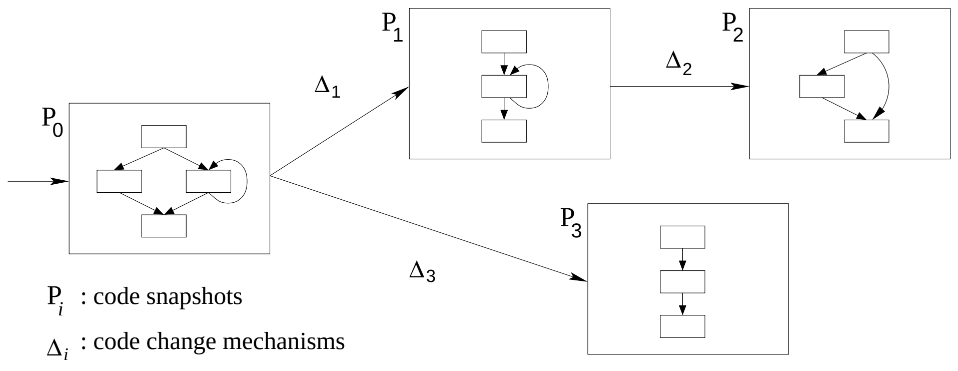

Figure 1. An example of a behavioral model for a self-modifying program

previously encountered—is to use dynamic analysis techniques, e.g., by running the binary in an emulator or under the control of a debugger. However, dynamic analysis can be tedious and time-consuming, and can be defeated via anti-monitoring and anti-virtualization techniques that allow the malware to detect when its code is being run under the control of a debugger or emulator [4, 6, 14, 19]. Furthermore, even if the unpacker code is activated during a dynamic analysis, the code may be unpacked over many rounds of unpacking in such a way that the amount of code that is materialized on any given round is small enough to escape detection, thereby leading to false negatives.

These shortcomings of dynamic malware analysis lead us to revisit the question of static analysis of malware binaries. However, this faces the hurdle that since self-unpacking malware effectively change their code on the fly, existing semantic bases for static analysis—which typically assume a fixed program—are inadequate for reasoning about them. We need a semantic framework that is able to capture the dynamic generation and modification of a program's code during execution. Furthermore, while most of the proposals for dynamic code modification in the research literature, e.g., in the context of just-in-time compilation [1] or dynamic code optimization [3, 16, 18], tend to modify code in principled ways, the code for dynamic malware unpacking tends to be sneaky and obscure, often using arbitrary arithmetic operations on memory such as addition, XOR, rotation, etc., to effect the code changes needed. This makes the development of a formal semantics for such programs a nontrivial challenge.

This paper makes two contributions in this regard. First, it takes a step in addressing this challenge by proposing a formal semantics for a simple "abstract assembly language" that permits dynamic code modification. Second, it shows a number of example applications for this semantic framework, including (1) static unpacking and analysis of packed malware binaries; (2) generalizing the traditional notions of "liveness" and "dead code" to self-modifying code, and applying this to malware deobfuscation; and (3) static identification of dynamic anti-monitoring defenses. Our ideas are not intended to replace the dynamic analyses currently used by security researchers, but to complement them and provide researchers with an additional tool for understanding malware.

2 Behavioral Models

Static analyses generally construct *behavioral models* for programs, i.e., static program representations that can be used to reason about its runtime behavior; examples include graph-based representations, such as control flow graphs and program dependence graphs, commonly used by compilers. Typically, such behavioral models contain program code (e.g., basic blocks) as well as control or data flow relationships between them that can be used to obtain insights into the execution behavior of the program.

Traditional program analyses typically assume that the program does not change during the course of execution. This assumption generally holds for ordinary application code, so it suffices to use a simple behavioral model that contains a single "snapshot" of a program's code. This assumption does not hold for malware, however, since malware code very often changes as execution progresses. For example, the Rustock rootkit and spambot arrives as a small unpacking routine together with an encrypted payload. When executed, this code unpacks the code for a rootkit loader, which subsequently carries out a second round of unpacking to generate the code for the rootkit itself [7]. This program, therefore, requires three different code snapshots to describe its behavior.

Since our goal is to reason statically about malware code, it becomes necessary for us to develop more general behavioral models that can cope with changes to the program code at runtime. Figure 1 shows an example this. It is a directed graph where each vertex P_i is a traditional static program representation and whose edges represent possible runtime changes: an edge $P_i \to P_j$ indicates that the code P_i may change to the code P_j at runtime. Furthermore, each edge has a label Δ_i that represents the mechanism for the corresponding code change. For example, each vertex may be a control flow graph for a code snapshot encountered at runtime while the label on each edge is a program slice that effects the corresponding code change. This allows us to reason about both the code snapshots P_i as they exist in between changes to the code, as well as the mechanisms Δ_i used to effect those changes. Sections 3 and 4 discuss a semantic formulation for machine-level programs that forms the underpinning of such reasoning; Section 5 then discusses the details of how such a behavioral model may be constructed.

3 An Abstract Assembly Language

In order to define and prove the correctness of static analyses for malware, we need to specify the language semantics underlying the analysis. This turns out to be problematic for malware because malware code is typically self-modifying, while most language semantics considered in the literature assume that the program code is immutable. We therefore begin by describing the syntax and semantics for a small, low-level abstract assembly language that allows dynamic code modification.

3.1 Syntax

Figure 2 gives the syntax of our abstract assembly language. For the sake of generality, we abstract away architecture-specific details, e.g., as the distinction between registers and memory, various addressing modes, etc. We assume an infinite set of memory locations indexed by the natural numbers **N**. Operations include assignment, input, and conditional and unconditional branches, the last of which include both direct and indirect branches. The set of instructions is denoted by **I**. Computations are over the set of natural numbers **N**.

Malware code is typically encountered in the form of executable binaries. We model this using the notion of memory maps: A *memory map* $M \in \mathbf{M} = \mathbf{N} \longrightarrow \mathbf{N}_\perp$ and denotes the contents of a program's memory (locations that do not have a defined value are assumed to be implicitly mapped to $\perp$).

Syntactic Categories:

a, n	$\in$	$\mathbf{N}$	(integers)
e	$\in$	$\mathbf{E}$	(expressions)
I	$\in$	$\mathbf{I}$	(instructions)
M	$\in$	$\mathcal{M} = \mathbf{N} \longrightarrow \mathbf{N}_\perp$	(memory map)
P	$\in$	$\mathcal{M} \times \mathbf{N}$	(programs)

Syntax:

e	::=	n	(constant)
	$\mid$	$\text{MEM}[e_1]$	(contents of memory)
	$\mid$	$e_1\ op\ e_2$	(arithmetic)
			$op \in \{+, -, \ldots\}$
I	::=	$\text{MEM}[e_1] := e_2$	(assignment)
	$\mid$	$\mathsf{input} \Rightarrow \text{MEM}[e_1]$	(input)
	$\mid$	$\mathsf{if}\ e_1\ \mathsf{goto}\ a$	(conditional branch)
	$\mid$	$\mathsf{goto}\ e$	(unconditional branch)
	$\mid$	halt	

Figure 2. An abstract assembly language: Syntax

An important aspect of low-level program representations on real processors is that there is no fundamental distinction between code and data: memory locations contain binary values that can be interpreted either as representing data or as encoding program instructions. This makes it possible to operate on a memory location as though it contains data, e.g., by adding or subtracting some value from it, and then interpret the result as code and execute it. To capture this, we assume the existence of an injective function $\mathsf{encode} : \mathbf{I} \longrightarrow \mathbf{N}$ such that, given $I \in \mathbf{I}$, the value $\mathsf{encode}(I) \in \mathbf{N}$ gives its binary representation. We have a corresponding notion of "decoding" a number to the instruction it represents, denoted by a function decode, which extends encode^{-1} to a total function by mapping numbers that do not correspond to legal instruction encodings to $\perp$:

$$\mathsf{decode}(n) : \mathbf{N} \longrightarrow I_\perp$$
$$\mathsf{decode}(n) = \begin{cases} I & \text{if } \exists I \in \mathbf{I} \text{ s.t. } \mathsf{encode}(I) = n \\ \perp & \text{otherwise} \end{cases}$$

Syntactically, a program $P \in \mathcal{M} \times \mathbf{N}$ consists of a memory map together with the entry point, i.e., the address where execution begins. Given a program $P = (M, a)$, the memory map M specifies the contents of all of P's memory, including both code and data.

To simplify the discussion, we assume that each instruction occupies a single location in memory. While this does not hold true of variable-instruction-length architectures such as the Intel IA-32, the issues raised by variable-length instructions are orthogonal to the topic of this paper.

3.2 Semantics

The semantics of this language are given in Figure 3. We use the following notation. The set $S_\perp$ denotes $S \cup \{\perp\}$. The Kleene closure of a set S, i.e., the set of all finite sequences of elements of S, is written S^*. Given a function $f : A \to B$, $f[a \mapsto b]$ denotes the function $\lambda x.[\text{if } x = a \text{ then } b \text{ else } f(x)]$, which is the same as f except at a, where its value is b.

A store $\sigma \in \mathcal{M}$ is simply a memory map, and maps locations to values. A program state (a, σ, θ) has three components: a is the "program counter value," which is either $\perp$ (indicating an undefined location), or $\bigcirc$ (indicating that the program is halted), or else is an integer giving the location of the next instruction to execute; $\sigma \in \mathcal{M}$ is the store and represents the contents of memory; and $\theta \in \Theta$ represents the unexpended input in the input stream. The equations defining the function $\mathscr{E}$, which specifies the meaning of expressions, are straightforward. The definition of $\mathscr{I}$, the semantic function describing the behavior of instructions, shows the low-level nature of the language: $\mathscr{I}$ takes a state as argument, uses the value of the program counter component a of its argument to retrieve the contents $\sigma(a)$ of the corresponding memory location, decodes this value using the decoding function decode, then executes the resulting instruction. The program counter value after the execution of an instruction at location a depends on the type of instruction: for control transfer instructions, the next program counter is specified as a target of the instruction itself; otherwise, the program counter value is that of the next instruction, at location $a + 1$.

Four of the five instructions in our language—assignment, conditional branches, unconditional branches, and halt—are familiar and have the expected semantics, as shown in Figure 3. The fifth instruction, '$\mathsf{input} \Rightarrow \text{MEM}[e]$', is used to read in external input. The idea here is that the next value in the input stream is written to the memory location with address e. We use the notion of "external input" to refer to any source that is external to the program: it may refer to a human user, another program executing on the same computer, input that is read in from the program's execution environment or over a network from a remote host, etc. (We note in passing that the language shown does not have instructions for procedure calls and returns. This omission is primarily to keep the discussion simple: it would be straightforward to extend our language to deal with procedure calls by adding call and return instructions to the instruction set together with an additional stack component to each state.)

Define the predicate hasSucc on states such that $\mathsf{hasSucc}(t)$ is true if and only if the state t has a successor, i.e., it is not undefined and not a halted state:

$$\mathsf{hasSucc}(t) \overset{\triangle}{=} t \neq \perp \text{ and } \forall \sigma, \theta : t \neq (\bigcirc, \sigma, \theta).$$

Given a program $P = (M, a_0)$ and an input stream θ, its initial state is given by $\sigma_0[P, \theta] = (M, a_0, \theta)$. Given this initial state, and the transition relation $\mathscr{I} : \Sigma \longrightarrow \Sigma$, we can now specify a trace semantics for a program as follows. First, define the set of finite traces of a program P for an input stream θ to be

$$\begin{aligned}\mathcal{T}[\![P, \theta]\!] = \{t0 \ldots t_k \mid\ & t_0 = \sigma_0[P, \theta] \text{ and} \\ & (\forall i : 0 \leq i < k)[\mathsf{hasSucc}(t_i) \text{ and} \\ & t_{i+1} = \mathscr{I}[\![t_i]\!]]\}.\end{aligned}$$

The set of all finite traces of a program P is then given by $\mathfrak{T}[\![P]\!] = \bigcup_{\theta \in \Theta} \mathcal{T}[\![P, \theta]\!]$.

3.3 Reachability

The notion of reachability will be important in the discussion that follows. Intuitively, a location ℓ_1 is statically reachable from a location ℓ_0 if, under the usual static analysis assumption that either branch of a conditional may be taken at runtime, it is possible for control to go from the instruction at location ℓ_0 to the instruction at location ℓ_1. The reason we refer to it as "*static* reachability" is that it assumes that the program code is static, i.e., is not being modified during execution. We first define the notion of a control-flow successor:

Definition 3.1 Let $P = (\sigma, a)$ be a program and $I = \mathsf{decode}(\sigma(\ell_0))$ the instruction in P at location ℓ_0. The location ℓ_1 is a *control-flow successor* of ℓ_0 in P (written '$\ell_0 \leadsto_P \ell_1$') if one of the following hold:

– $I = \ulcorner \text{MEM}[e_1] := e_2 \urcorner$ and $\ell_1 = \ell_0 + 1$;

– $I = \ulcorner \mathsf{input} \Rightarrow \text{MEM}[e_1] \urcorner$ and $\ell_1 = \ell_0 + 1$;

– $I = \ulcorner \mathsf{if}\ e_1\ \mathsf{goto}\ a_1 \urcorner$ and $\ell_1 \in \{a_1, \ell_0 + 1\}$;

– $I = \ulcorner \mathsf{goto}\ e_1 \urcorner$ and $\ell_1 = \mathscr{E}[\![e_1]\!]\sigma$.

▮

Value Domains:

$$\begin{array}{rcll}
\sigma \in \mathcal{M} & = & \mathbf{N} \longrightarrow \mathbf{N}_\perp & \text{(store)} \\
\theta \in \Theta & = & \mathbf{N}^* & \text{(input stream)} \\
\pi & = & \mathbf{N} \cup \{\perp, \bigcirc\} & \text{(program counter values)} \\
\Sigma & = & \pi \times \mathcal{M} \times \Theta & \text{(program states)}
\end{array}$$

Semantics:

EXPRESSIONS:

$$\begin{array}{lcl}
\mathscr{E} : \mathbf{E} \times \mathcal{M} \longrightarrow \mathbf{N}_\perp & & \\
\mathscr{E}\,[\![n]\!]\,\sigma & = & n \\
\mathscr{E}\,[\![\mathrm{MEM}[e]]\!]\,\sigma & = & \textbf{let } n = \mathscr{E}\,[\![e]\!]\,\sigma \textbf{ in} \\
 & & \quad \textbf{if } n \neq \perp \textbf{ then } \sigma(n) \\
 & & \quad \textbf{else } \perp \\
\mathscr{E}\,[\![e_1 \ op\ e_2]\!]\,\sigma & = & \textbf{let } n_1 = \mathscr{E}\,[\![e_1]\!]\,\sigma \ n_2 = \mathscr{E}\,[\![e_2]\!]\,\sigma \textbf{ in} \\
 & & \quad \textbf{if } n_1 \neq \perp \text{ and } n_2 \neq \perp \textbf{ then } n_1 \ op\ n_2 \\
 & & \quad \textbf{else } \perp
\end{array}$$

INSTRUCTIONS: The semantic function $\mathscr{I} : \Sigma \rightarrow \Sigma$ gives the effect of executing an instruction. It effectively specifies the transition relation between states.

$$\begin{array}{lcl}
\mathscr{I}([\![a]\!], \sigma, \theta) \ = \ \textbf{case } \mathsf{decode}(\sigma(a)) \textbf{ of} & & \\
\quad \ulcorner \mathrm{MEM}[e_1] := e_2 \urcorner & : & \textbf{let } n_1 = \mathscr{E}\,[\![e_1]\!]\,\sigma, n_2 = \mathscr{E}\,[\![e_2]\!]\,\sigma \textbf{ in} \\
 & & \quad \textbf{if } n_1 \neq \perp \text{ and } n_2 \neq \perp \textbf{ then } (a+1, \sigma[n_1 \mapsto n_2], \theta)\textbf{; else } \perp \\
\quad \ulcorner \mathsf{input} \Rightarrow \mathrm{MEM}[e_1] \urcorner & : & \textbf{let } n_1 = \mathscr{E}\,[\![e_1]\!]\,\sigma \textbf{ in} \\
 & & \quad \textbf{if } n_1 \neq \perp \text{ and } \mathit{length}(\theta) > 0 \textbf{ then } (a+1, \sigma[n_1 \mapsto \mathit{hd}(\theta)], \mathit{tl}(\theta))\textbf{; else } \perp \\
\quad \ulcorner \mathsf{if}\ e_1\ \mathsf{goto}\ a_1 \urcorner & : & \textbf{let } n_1 = \mathscr{E}\,[\![e_1]\!]\,\sigma \textbf{ in} \\
 & & \quad \textbf{if } n_1 \neq \perp \textbf{ then} \\
 & & \quad\quad \textbf{if } n_1 \neq 0 \textbf{ then } (a_1, \sigma, \theta)\textbf{; else } (a+1, \sigma, \theta) \\
 & & \quad \textbf{else } \perp \\
\quad \ulcorner \mathsf{goto}\ e_1 \urcorner & : & \textbf{let } n_1 = \mathscr{E}\,[\![e_1]\!]\,\sigma \textbf{ in} \\
 & & \quad \textbf{if } n_1 \neq \perp \textbf{ then } (n_1, \sigma, \theta)\textbf{; else } \perp \\
\quad \ulcorner \mathsf{halt} \urcorner & : & (\bigcirc, \sigma, \theta) \\
\quad \perp & : & \perp
\end{array}$$

Figure 3. An abstract assembly language: Semantics

The first two cases above correspond to execution fall-through in non-control-transfer instructions, while the remaining two cases refer to the behavior of control transfers. The notion of static reachability is simply the reflexive transitive closure of the control-flow successor relation:

Definition 3.2 A location ℓ_1 is *statically reachable* from a location ℓ_0 in a program P, denoted by $\ell_0 \rightsquigarrow_P^* \ell_1$, if: (i) $\ell_0 = \ell_1$; or (ii) there exists a location ℓ such that $\ell_0 \rightsquigarrow_P \ell$ and $\ell \rightsquigarrow_P^* \ell_1$. ▮

4 Giving Structure to Dynamic Code Modification

Before we can discuss dynamically modified code, we must specify exactly what that term means. We interpret this narrowly. It is not enough for the program to simply modify some memory regions that typically contain code,[2] such as the *.text* section, since programs sometimes contain data embedded in the instruction stream. On the other hand, we want to be able to catch programs that dynamically generate and execute code in memory regions that usually do not contain code, e.g., the stack or heap. We require, therefore, that a memory location must be modified (i.e., written to) and subsequently executed:

Definition 4.1 A trace $T = t_0 \dots t_i \dots t_j \dots$ is code-modifying if and only if it contains states t_i and t_j, $j > i$, such that t_i modifies a memory location a and t_j executes the instruction at location a.

A program P is code-modifying if and only if there is some trace $T \in \mathfrak{T}\,[\![P]\!]$ that is code-modifying. ▮

[2]Some researchers have used this criterion as a heuristic to identify self-modifying code [17].

A particular execution of a program can modify its code numerous times [7]. Given an execution trace T for a program P, it is conceptually convenient to distinguish between the code that is changing memory locations that may be executed at some future point from the code that is encountered when one of those modified locations is eventually executed. We do this by dividing up an execution trace into *phases*:

Definition 4.2 A phase T' of a trace T is a maximal subsequence of T such that T' does not execute any location modified by T'. ▮

This means that all phases but the first always begin with the execution of an instruction that was created and/or modified in a previous phase.

The following result captures a crucial property of phases: namely, that for each state S_i in the phase, the contents of memory at the program counter location a_i for that state is the same as the contents of location a_i at the beginning of that phase.

Theorem 4.1 *Let $\phi = S_0, \ldots, S_i, \ldots$ be a phase, with $S_i = (a_i, M_i, \theta_i)$, then $M_i(a_i) = M_0(a_i)$ for all $S_i \in \phi$.*

Proof By contradiction. Suppose that the proposition does not hold, i.e., $M_i(a_i) \neq M_0(a_i)$ for some state $S_i \in \phi$. Then it must be the case that memory location a_i is modified by some earlier state $S_j \in \phi$, $j < i$. Thus, $S_i \in \phi$ executes a location modified by $S_j \in \phi$. From the definition of a phase, this means that S_i and S_j cannot be in the same phase. Thus, either $S_i \notin \phi$ or $S_j \notin \phi$, which is a contradiction. The theorem follows. ▮

This theorem implies that for each state S_i in a phase, the instruction "seen" at location a_i by S_i, when control reaches a_i, is the same as the instruction at that location at the beginning of the phase:

Corollary 4.2 *Let $\phi = S_0, \ldots, S_i, \ldots$ be a phase for a program, with $S_i = (a_i, M_i, \theta_i)$. Then for all $S_i \in \phi$, the instruction at location a_i that is executed in state S_i is the same as the instruction occurring at location a_i in the initial state S_0 of ϕ.*

This corollary means that the static code "snapshot" at the beginning of a phase gives the program that is executed by that phase, and can be used as a basis for analyzing and understanding the behavior of that phase.

We next consider the question of partitioning a trace into phases. Intuitively, given a trace of a program's execution, we can simply replay the trace keeping track of memory locations that are modified due to the execution of each instruction, and thereby identify those states where the instruction about to be executed is at a location that was modified in the current phase. Each such state then marks the beginning of a new phase. More formally, the process of partitioning a trace into phases can be described as follows. For any instruction I, let $\textsc{mod}\,[\![I]\!]\,\sigma$ denote the set of locations modified by I given a store σ:

$$\textsc{mod}\,[\![I]\!]\,\sigma = \begin{cases} \{n_1\} & \text{if } I = \ulcorner \textsc{mem}[e_1] := e_2 \urcorner \\ & \text{and } n_1 = \mathscr{E}\,[\![e_1]\!]\,\sigma \\ \{n_1\} & \text{if } I = \ulcorner \mathsf{input} \Rightarrow \textsc{mem}[e_1] \urcorner \\ & \text{and } n_1 = \mathscr{E}\,[\![e_1]\!]\,\sigma \\ \emptyset & \text{otherwise} \end{cases}$$

Given a trace $T = t_0 t_1 \ldots t_n$, let $t_i = (a_i, \sigma_i, \theta_i)$, let the instruction at location a_i, which is about to be executed, be $I_i = \mathsf{decode}(\sigma(a_i))$. We first define functions Δ^- and Δ^+ that give the set of memory side effects of the current phase: $\Delta^-(t)$ gives the set of locations written to upto state t, i.e., at the point where the instruction I specified by t's program counter is about to be executed, while $\Delta^+(t)$ gives this set after I has executed. For any state $t = (a, \sigma, \theta)$, let $I = \mathsf{decode}(\sigma(a))$ be the instruction executed in t, then the value of $\Delta^+(t)$ is defined as follows:

$$\Delta^+(t) = \Delta^-(t) \cup \textsc{mod}\,[\![I]\!]\,\sigma$$

The function Δ^- is defined as follows. For the very first state t_0 of the trace, no code has been executed and therefore no memory modifications have occurred. For a later state t_{i+1}, the value of $\Delta^-(t_{i+1})$ depends on whether t_{i+1} is the first state in a new phase. If the program counter a_{i+1} in this state is an address that has been modified in $\Delta^+(t_i)$, then t_{i+1} starts a new phase, and since no instructions in this new phase have been executed, the memory effects of the phase are empty. Otherwise, t_{i+1} simply extends the current phase, so $\Delta^-(t_{i+1}) = \Delta^+(t_i)$. Thus, we have:

$$\begin{aligned} \Delta^-(t_0) &= \emptyset \\ \Delta^-(t_{i+1}) &= \begin{cases} \emptyset & \text{if } a_{i+1} \in \Delta^+(t_i) \\ \Delta^+(t_i) & \text{otherwise} \end{cases} \end{aligned}$$

We can now specify how to partition a trace T into a sequence of phases. Given a state $t \in T$, let $\phi(t) \in \mathbf{N}$ be its *phase number*, which denotes which phase it belongs to. We can use the Δ^+ function to assign phase numbers to states:

$$\phi(t_0) = 0$$
$$\phi(t_{i+1}) = \begin{cases} \phi(t_i) + 1 & \text{if } a_{i+1} \in \Delta^+(t_i) \\ \phi(t_i) & \text{otherwise} \end{cases}$$

Then, each phase ϕ_i is a maximal subsequence of T such that all of the states in ϕ_i have the same phase number.

Recall that, as mentioned in Section 2, one of our goals is to construct static behavioral models for programs with dynamically modified code, which capture both "code snapshots" as well as code change mechanisms, i.e., the code that is responsible for dynamically modifying the code. In preparation for this, we first outline how we can abstract away from the trace semantics of a program to a description that records only the possible code snapshots and code change mechanisms that occur over all possible executions of the program. First, given a phase $\phi = t_0, \ldots, t_n$ in a trace T, where $t_i = (a_i, \sigma_i, \theta_i)$, $0 \leq i \leq n$, we specify two characteristics of ϕ:

– The code snapshot for the phase, denoted by $\mathsf{code}(\phi)$. From Theorem 4.1, this is given by the program at the first state of the phase:

$$\mathsf{code}(\phi) = (\sigma_0, a_0).$$

– The code responsible for changes to memory that occur during the execution of ϕ, denoted by $\mathsf{dmods}(\phi)$. Since the set of memory locations modified during the execution of ϕ is given by $\Delta^+(t_n)$, this is given by the dynamic slice of $\mathsf{code}(\phi)$ for the set of locations (i.e., variables) $\Delta^+(t_n)$ and the history ϕ.

We can now define the *dynamic evolution graph* for a program P to be an edge-labelled directed graph with vertices V and edges E, given by the following:

– Let $T \in \mathfrak{T}[\![P]\!]$ be any execution trace for P, and let ϕ be any phase of T. Then $\mathsf{code}(\phi)$ is a vertex in V.

– Let $T \in \mathfrak{T}[\![P]\!]$ be any execution trace for P; ϕ_i and ϕ_{i+1} be any two successive phases of T; and $D = \mathsf{dmods}(\phi_i)$. Let v_i and v_{i+1} be vertices in V corresponding to phases ϕ_i and ϕ_{i+1} respectively. Then $v_i \xrightarrow{D} v_{i+1}$ is an edge in E.

The next section discusses how we can statically approximate a program's dynamic evolution graph to construct a behavioral model which we call its "program evolution graph."

5 Program Evolution Graphs

This section discusses one concrete application of the semantics described in the previous section. We show how this semantics can be used as a formal basis for constructing static behavioral models for self-modifying programs, which can then be used to reason about the behavior of such programs.

The dynamic evolution graph for a program P, discussed in the previous section, is an abstraction of its trace semantics that describes some aspects of P's runtime behavior. The *program evolution graph* for P is a static approximation to P's dynamic evolution graph. This section discusses the construction of such graphs.

Given an executable P, we initialize the program evolution graph of G to contain P as its only vertex, then proceed by iteratively analyzing the vertices. For each vertex, our algorithm identifies *transition edges* that specify control transfers into newly modified code (Section 5.1), then processes each transition edge separately to obtain (an approximation to) the code that would result from dynamic code modifications leading up to that instruction ("static unpacking," Section 5.2). The input executable P has to contain the code for the initial unpacker (which unpacker cannot itself be in packed form), as well as a control flow path to this code so that it can be executed; our algorithm begins by discovering transition edges out of this initial unpacker, then iteratively processes the unpacked programs it identifies. The resulting programs from static unpacking are then added to the program evolution graph and eventually processed in their turn. The discussion below focuses on the core computations of of this nested iteration, namely, identification of transition points and the static unpacking for any given transition point. The overall algorithm is summarized in Figure 4.

5.1 Identifying Transition Edges

The first step of the analysis involves identifying those instructions (equivalently, locations) in a program that may mark a boundary between phases, i.e., from which control may reach a location that has been modified by the current phase. We refer to such control flow as a *transition edge*. For determining the locations that may be modified during execution, we assume a binary-level alias analysis that gives a superset of the set of possible referents of all (direct and indirect) memory references. The issue of alias analysis—and, in particular, binary-level alias analysis [12, 13]—is well-studied and in any case orthogonal to the topic of this paper, and we do not pursue the mechanics of such analyses further

Input: A program $P_0 = (M_0, a_0)$.

Output: A program evolution graph $G = (V, E)$.

Method:

```
$V = \{P_0\}$;    /* $P_0$ is unmarked */
$E = \emptyset$;
while there are unmarked elements of $V$ do
    let $P = (\sigma, a)$ be any unmarked element of $V$;
    mark $P$;
    $T = \mathsf{TransitionEdges}(P)$;                /* see Section 5.1 */
    for each $t = (b, c) \in T$ do                  /* static unpacking: see Section 5.2 */
        $M := Mods_P(a, b)$;                        /* locations modified on paths from $a$ to $b$
        $S :=$ backward static slice of $P$ at $b$ with respect to $M$;
        $\sigma' :=$ store resulting from the symbolic execution of $S$;
        $P' := (\sigma', c)$;                       /* unpacked program */
        if $P' \notin V$ then
            add $P'$ (unmarked) to $V$; add '$P \xrightarrow{S} P'$ to $E$;
        fi
    od    /* for */
od    /* while */
return $G$;
```

Figure 4. Algorithm for static construction of program evolution graphs

here. We use the results of the alias analysis to identify transition edges.

More formally, given a location ℓ and an instruction I, let $alias_I(\ell)$ denote (a superset of) the set of values of memory location ℓ when control reaches I over all possible executions of the program.[3] We generalize this to $val_I(e)$, denoting the set of possible values taken on by any expression e when control reaches I, by lifting pointwise: the details are given in Figure 5. Then, for each instruction I we compute two sets: $write(I)$, the set of memory locations that may be written to by I; and $next(I)$, the set of locations that control may go to after the execution of I. The details of how these functions are defined are given in Figure 5. Note that $next(I)$ is in fact simply a static approximation to the semantic relation $\leadsto_P$ defined in Section 3.3. Transition edges are then obtained as follows:

Definition 5.1 A transition edge for a program P is a pair of locations (a, b) such that the following hold:

1. $a \notin write(X)$ for any instruction $X \in P$; and
2. let the instruction at location a be I_a, then there exists an instruction J in the program, at location ℓ_J, such that a is reachable from ℓ_J and $b \in write(J) \cap next(I_a)$.

∎

The definition states that while location a is not modified by the program (condition 1), location b may be modified by J and control can then go from J to a and thence to b (condition 2). Control enters the modified code when it goes from a to b.

In general, a program may have multiple transition edges, corresponding to different ways of modifying a program. Given a program P, we denote its set of transition edges—which will, in general, be dependent on the precision of the alias analysis used—by $\mathsf{TransitionEdges}(P)$.

5.2 Static Unpacking

Once we have the set of transition edges for a program P, we process them individually to determine, as far as possible, the contents of memory after unpacking.

[3]In the context of more familiar source-level analyses, pointer alias analyses compute, for each pointer, the set of objects it may point to. At the assembly-code or machine-code level, this is equivalent to computing, for each location that may be used as a pointer, the set of possible addresses—i.e., the values—contained at that location. In our case, therefore, alias analysis amounts to computing the possible values at each location in memory.

Given: an instruction I at location ℓ_I.

$val_I(e)$ denotes the set of possible values of an expression e when control reaches I for all possible executions of the program:

$$val_I(e) = \begin{cases} \{n\} & \text{if } e = \ulcorner n \urcorner \\ \bigcup\{alias_I(e') \mid e' \in val_I(e_1)\} & \text{if } e = \ulcorner \text{MEM}[e_1] \urcorner \\ \{e'_1 \; op \; e'_2 \mid e'_1 \in val_I(e_1) \text{ and } e'_2 \in val_I(e_2)\} & \text{if } e = \ulcorner e_1 \; op \; e_2 \urcorner \end{cases}$$

$write(I)$ denotes the set of locations that may be modified by I over all possible executions of the program:

$$write(I) = \begin{cases} val_I(e_1) & \text{if } I = \ulcorner \text{MEM}[e_1] := e_2 \urcorner \\ val_I(e_1) & \text{if } I = \ulcorner \mathsf{input} \Rightarrow \text{MEM}[e_1] \urcorner \\ \emptyset & \text{otherwise} \end{cases}$$

$next(I)$ denotes the set of locations that control may go to immediately after the execution of I:

$$next(I) = \begin{cases} \{\ell_I + 1\} & \text{if } I = \ulcorner \text{MEM}[e_1] := e_2 \urcorner \\ \{\ell_I + 1\} & \text{if } I = \ulcorner \mathsf{input} \Rightarrow \text{MEM}[e_1] \urcorner \\ \{a, \ell_I + 1\} & \text{if } I = \ulcorner \mathsf{if} \; e_1 \; \mathsf{goto} \; a \urcorner \\ val_I(e) & \text{if } I = \ulcorner \mathsf{goto} \; e \urcorner \\ \emptyset & \text{if } I = \ulcorner \mathsf{halt} \urcorner \end{cases}$$

Figure 5. The functions *val*, *write*, and *next* (see Section 5.1)

We refer to this process as static unpacking. Of course, given the usual undecidability problems that arise during static analysis, this will not always be possible in general, and our analysis will sometimes conservatively indicate the contents of some memory locations to be unknown.

Given a transition edge (b, c) in a program $P = (\sigma, a)$, we first compute the set of locations $Mods_P(a, b)$ that may be modified during the execution of P along paths from a to b:

$$Mods_P(a, b) = \bigcup_{I \in P} \{write(I) \mid a \rightsquigarrow_P^* I \text{ and } I \rightsquigarrow_P^* b\}.$$

Our next step is to isolate the fragment of P whose execution may affect the values of locations in $Mods_P(a, b)$. For this, we compute S, the backward static slice of P for the set of locations (i.e., variables) $Mods_P(a, b)$. The store resulting from these modifications—which includes the modified code—is then obtained by symbolic execution of S.

This symbolic execution of the slice S begins in the memory state σ given by the initial store σ and proceeds until it reaches the point where control would go from b to c (recall that (b, c) is the transition edge under consideration). The details of exactly how the symbolic execution are carried out involve design and representation decisions concerning, for example, tradeoffs between cost and precision. For this reason, we do not pursue these details here.

5.3 Putting it all Together

After static unpacking of each transition edge, we update the program evolution graph. Given a program $P = (\sigma, a)$ and a transition edge (b, c), let the memory state resulting from static unpacking be σ', then the result of static unpacking is the program $P' = (\sigma', c)$. If P' is new, we add it to the set of vertices of the program evolution graph and the edge '$P \xrightarrow{S} P'$' to the set of its edges.

Since one of the main reasons for constructing a program evolution graph is for the analysis and understanding of malware code, we may also want to disassemble the program at each of its vertices and construct a control flow graph for it. Given a vertex $P = (\sigma, a)$, the locations to be disassembled are given by $\{\ell \mid a \rightsquigarrow_P^* \ell\}$.

The control flow graph for the resulting instruction sequence can be constructed in the usual way.

6 Applications of Program Evolution Graphs

This section sketches some applications of static analysis of program evolution graphs for understanding the behavior of malware binaries.

6.1 Exec-Liveness and Malware Deobfuscation

The discussion of static unpacking in Section 5.2 is quite liberal in the set of memory modifications it considers: given a program $P = (\sigma, a)$ and a transition edge (b, c), all memory modifications possible along all paths from a to b are considered as potentially part of the unpacked code. This may seem overly conservative, since it is entirely possible that some of these writes to memory are modifying data rather than code, or may even be "junk" writes intended to obfuscate. We include them all when computing the backward static slice used for static unpacking, nevertheless, since we do not know, ahead of time, which locations may be executed at some future point in the computation.

Once we have the entire program evolution graph, we can use static analysis to identify, at each vertex, locations that may be executed, without first being modified, along some path from that vertex through the program evolution graph. The idea is very similar to that of liveness analysis, and we refer to it as *exec-liveness*. We compute exec-liveness as follows. Given a vertex $v = (\sigma, a)$ in a program evolution graph, define the following sets for v:

- $write(v) = \bigcup\{write(I) \mid I$ is an instruction in $v\}$. This gives the set of locations that may be modified by the execution of v.
- $exec(v) = \{\ell \mid a \rightsquigarrow_v^* \ell\}$. This gives the set of code addresss that are statically reachable from the entry point a of the program v.

Let $\mathsf{xLive}_{in}(v)$ and $\mathsf{xLive}_{out}(v)$ denote, respectively, the sets of exec-live locations at entry to, and exit from, vertex v. These can be defined by the following dataflow equations:

$$\begin{aligned} \mathsf{xLive}_{in}(v) &= (\mathsf{xLive}_{out}(v) - write(v)) \cup exec(v) \\ \mathsf{xLive}_{out}(v) &= \bigcup \{\mathsf{xLive}_{in}(v') \mid v' \text{ a successor of } v\} \end{aligned}$$

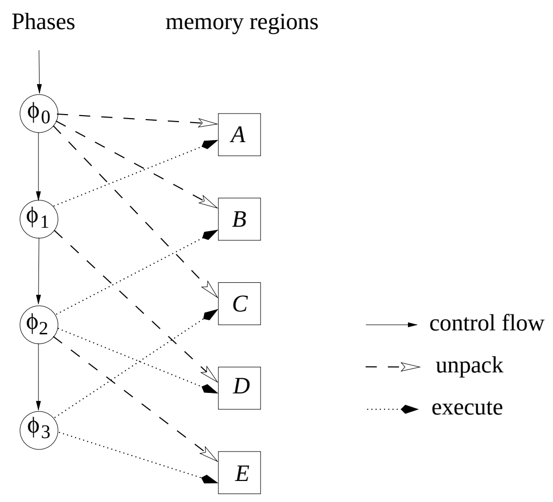

Figure 6. Way-Ahead-of-Time Unpacking

The boundary conditions for these equations are given by vertices that have no successors: the xLive_{out} sets for such vertices is $\emptyset$. These equations can be solved iteratively to obtain, for any vertex in the program evolution graph, the set of locations that may be executed in a later phase.

This information can then be used to simplify the program evolution graph. The essential idea is that a memory location ℓ is *relevant* at a program point p if it is live or exec-live at p, or if there is a path from p through the program evolution graph along which the value of location ℓ may be used before it is redefined; once a program evolution graph has been constructed, the computation of relevant locations is a straightforward application of dataflow analysis techniques [2]. Once such relevant locations have been identified in this way, assignments to irrelevant locations can be eliminated. The notion of relevance generalizes the traditional compiler notion of liveness, and irrelevant-code elimination generalizes a common compiler optimization known as dead code elimination. The difference here is that, instead of improving performance, the intent of irrelevant-code elimination is to remove useless instructions inserted for obfuscation purposes and make the program easier to understand and analyse (Christodorescu *et al.* refer to such assignmens as "semantic no-ops" [9]).

6.2 "Way Ahead of Time" Unpacking

One intuitively expects a program to unpack some code, then execute the unpacked code, perhaps unpack some more code and execute it, and so on—i.e., have each phase unpack only the code needed for the next phase. While intuitively straightforward, such a scheme is not general enough to capture all possible unpacking behaviors. In particular, it is possible for a phase to unpack code that is not used until a phase later than the next one. This is illustrated by Figure 6, where dashed lines indicate which memory regions are unpacked by a phase while dotted lines show the memory regions a phase executes. Phase ϕ_0 unpacks three memory regions: A, B, and C. The next phase, ϕ_1, executes only one of these (region A), and unpacks a fourth region, D. The following phase, ϕ_2, executes regions B (unpacked by ϕ_0) and D (unpacked by ϕ_1), and unpacks region E. Finally, phase ϕ_3 executes region C (unpacked by ϕ_0) and E (unpacked by ϕ_2). In this example, phase ϕ_0 unpacks certain regions of memory that are executed, not by the next phase ϕ_1, but rather by subsequent phases (ϕ_2 and ϕ_3). We refer to this kind of unpacking as "way-ahead-of-time (WAT) unpacking."

WAT-unpacking can be problematic for dynamic malware analysis tools since it effectively requires that the tools maintain a complete history of a program's unpacking behavior, which can be expensive due to the additional memory requirements. In practice, such tools sometimes make the simplifying assumption that each phase only unpacks code that is used by the next phase, i.e., that WAT-unpacking does not take place. This can compromise soundness by causing some unpacked code to fail to be detected, as in the case of the Renovo dynamic analysis tool [15]. On the other hand, if a static analysis can guarantee the absence of WAT-unpacking in a malware executable, subsequent dynamic analyses of that executable can safely ignore its unpacking history, thereby improving efficiency significantly without losing soundness.

WAT-unpacking can be detected by static analysis of program evolution graphs via a straightforward adaptation of the notion of def-use chains [2]. Using the *exec* and *write* sets for the vertices of a program evolution graph, we can construct "def-exec chains," which are exactly analogous to the traditional static analysis notion of def-use chains and links each modification of a location (obtainable from the *write* set of each vertex) to all of its "uses" (obtainable from the *exec* set of each vertex). It is then straightforward to determine whether a location modified by one phase (vertex) v_1 in a program evolution graph is used by some other phase (vertex) v_2 where v_2 is not an immediate successor of v_1.

6.3 Conditionally Activated Unpacking

There are situations where either the unpacker or the unpacked code may be activated conditionally based on the value of some input. Examples of such behavior include the use of anti-monitoring and anti-virtualization defenses [4, 6, 14, 19], where the malware does not reveal itself if it determines that its execution is being watched; "time bombs" and "logic bombs," where the activation occurs only on specific dates (time bombs) or due to some specific environmental trigger (logic bombs); and bots that activate specific components of their payload only when they receive the appropriate command from a botmaster [10]. We can use static analysis to identify such situations:

Proposition 6.1 *Given a program P containing some dynamic unpacking code C and a conditional branch B, the dynamic unpacker code C is conditionally activated by B if the following hold:*

1. *the test expression e is computed from a value read in by an input instruction '$I \equiv$ '`input` $\Rightarrow \text{MEM}[e]$';*
2. *one of the targets of the branch instruction B leads to C; and*
3. *the other target of B does not lead to C without first going through the input instruction I.*

The reason for the third condition is that some malware, such as a bot, may simply sit in a loop, repeatedly reading in commands and checking to see if any of them direct it to activate the unpacker code.

The detection of dynamic defenses in this manner can give security researchers information about the location and nature of conditionally activated unpackers, thereby allowing them to take countermeasures if they choose to use dynamic analysis.

6.4 Cross-Phase Code Analysis

A number of researchers have described techniques for identifying malware based on examining its code for incriminating patterns, actions or behaviors. Most modern anti-virus products use "generic decryptors" to deal with polymorphic viruses. The idea here is to emulate the malware code for some amount of time, with the expectation that this will allow enough of the malware payload to become unpacked that it can be detected simply by scanning the contents of memory for telltale byte sequences [20]. Christodorescu *et al.* propose a more general scheme that uses instruction sequence templates to specify certain kinds of malicious behavior [8]. A significant drawback with a great deal of such work is that they assume that "enough" of the code of the candidate binary will be available at one time for analysis. This assumption can make it possible for a malware instance to evade dynamic security analysis tools by taking its actions and distributing them across multiple unpacking phases, so that no single phase exposes very much of the malware code.

These problems can be mitigated by static analysis of malware code where the entire program evolution graph is available for analysis. As a simple example, the malware detection approach of Christodorescu *et al.* [8] can be extended to deal with multiple unpacking phases as follows:

> *Given an executable program P with program evolution graph $G = (V, E)$, a malware template τ matches P if either of the following hold:*
>
> 1. *there is some vertex $v \in V$ such that τ matches v; or*
> 2. *there exists a path $v_1, \ldots, v_n$ in G such that τ matches the result of concatenating the code sequences for $v_1, \ldots, v_n$.*

7 Related Work

The only other work we are aware of that gives a formal semantics for self-modifying code is a recent paper by Cai *et al.* [5]. Their goals, namely, verification of code that is self-modifying but not in a "bad" way, are very different from ours, namely, static analysis of programs that may be malicious. The two approaches also differ considerably in their technical details: Cai *et al.* describe an approach based on Hoare logic that they use to reason about self-modifying code, while our approach is based on a trace semantics.

Dalla Preda *et al.* have discussed the use of abstract interpretation, based on a low-level trace semantics, to reason about low-level code obfuscations used by malware [11]. They do not focus on dynamic code modifications, but assume instead that the program has been unpacked and correctly disassembled and is available for analysis; their goal is to use abstract interpretation to hide the effects of code obfuscations used by malware, and to use such abstract interpretations to reason about the soundness and completeness of malware detectors. By contrast, we focus on the actual semantics of self-modifying code; our goal is to use this semantics as a basis for constructing behavioral models for, and reasoning about, self-modifying programs. The two works are orthogonal in the sense that the behavioral models constructed using the approach described here could be used as input to the techniques described by Dalla Preda *et al.*.

Christodorescu *et al.* have proposed incorporating some knowledge of instruction semantics into virus scanners to get around simple code obfuscations used by malware [8]. The semantic information they propose to use is much simpler than that presented here, and is based on a notion of "templates," which are instruction sequences with place-holders for registers and constants. They do not consider the semantic issues raised by code self-modification, but assume that the malware has been unpacked and is available for analysis. This work is similarly orthogonal to that described here, and (as discussed in Section 6.4) could be generalized using our behavioral models.

There is a large body of literature on malware detection and analysis; Szor gives a comprehensive summary [20]. These approaches typically use dynamic analysis, e.g., via generic decryptors, to deal with runtime code modification.

8 Conclusions

The rapid growth in the incidence of malware has made their detection and analysis of fundamental importance in computer system security. Most malware instances encountered today use packed (i.e., encrypted or compressed) payloads in an attempt to avoid detection, unpacking the payloads at runtime via dynamic code modification. Modern malware detectors use dynamic analysis to get around this, but this approach has the drawback that it may not be able to detect conditionally unpacked code, e.g., due to anti-debugging defenses in the malware or because the malware payload is activated only on a specific date or time or in response to some specific command.

This paper describes an approach that complements such dynamic techniques via static analysis of malware code. We present a formal semantics that can cope with self-modifying code, then use this to develop the notion of a phase-based concrete semantics that describes how a program's code evolves due to dynamic code modifications. We then discuss how this concrete semantics can be statically approximated. Our ideas can be used to understand some aspects of the behavior of dynamically modified code, e.g., the use of anti-monitoring defenses, and also to statically approximate the code that is unpacked at runtime.

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