More Enforceable Security Policies

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Abstract

We analyze the space of security policies that can be enforced by monitoring programs at runtime. Our program monitors are automata that examine the sequence of program actions and transform the sequence when it deviates from the specified policy. The simplest such automaton truncates the action sequence by terminating a program. Such automata are commonly known as security automata, and they enforce Schneider's EM class of security policies. We define automata with more powerful transformational abilities, including the ability to insert a sequence of actions into the event stream and to suppress actions in the event stream without terminating the program. We give a set-theoretic characterization of the policies these new automata are able to enforce and show that they are a superset of the EM policies.

1 Introduction

When designing a secure, extensible system such as an operating system that allows applications to download code into the kernel or a database that allows users to submit their own optimized queries, we must ask two important questions.

1. What sorts of security policies can and should we demand of our system?
2. What mechanisms should we implement to enforce these policies?

Neither of these questions can be answered effectively without understanding the space of enforceable security policies and the power of various enforcement mechanisms.

Recently, Schneider [Sch00] attacked this question by defining EM, a subset of safety properties [Lam85, AS87] that has a general-purpose enforcement

mechanism - a *security automaton* that interposes itself between the program and the machine on which the program runs. It examines the sequence of security-relevant program actions one at a time and if the automaton recognizes an action that will violate its policy, it terminates the program. The mechanism is very general since decisions about whether or not to terminate the program can depend upon the entire history of the program execution. However, since the automaton is only able to recognize bad sequences of actions and then terminate the program, it can only enforce safety properties.

In this paper, we re-examine the question of which security policies can be enforced at runtime by monitoring program actions. Following Schneider, we use automata theory as the basis for our analysis of enforceable security policies. However, we take the novel approach that these automata are transformers on the program action stream, rather than simple recognizers. This viewpoint leads us to define two new enforcement mechanisms: an *insertion automaton* that is able to insert a sequence of actions into the program action stream, and a *suppression automaton* that suppresses certain program actions rather than terminating the program outright. When joined, the insertion automaton and suppression automaton become an *edit automaton*. We characterize the class of security policies that can be enforced by each sort of automata and provide examples of important security policies that lie in the new classes and outside the class EM.

Schneider is cognizant that the power of his automata is limited by the fact that they can only terminate programs and may not modify them. However, to the best of our knowledge, neither he nor anyone else has formally investigated the power of a broader class of runtime enforcement mechanisms that explicitly manipulate the program action stream. Erlingsson and Schneider [UES99] have implemented inline reference monitors, which allow arbitrary code to be executed in response to a violation of the security policy, and have demonstrated their effectiveness on a range of security policies of different levels of abstraction from the Software Fault Isolation policy for the Pentium IA32 architecture to the Java stack inspection policy for Sun's JVM [UES00]. Evans and Twyman [ET99] have implemented a very general enforcement mechanism for Java that allows system designers to write arbitrary code to enforce security policies. Such mechanisms may be more powerful than those that we propose here; these mechanisms, however, have no formal semantics, and there has been no analysis of the classes of policies that they enforce. Other researchers have investigated optimization techniques for security automata [CF00, Thi01], certification of programs instrumented with security checks [Wal00] and the use of run-time monitoring and checking in distributed [SS98] and real-time systems [KVBA+99].

Overview The remainder of the paper begins with a review of Alpern and Schneider's framework for understanding the behavior of software systems [AS87, Sch00] (Section 2) and an explanation of the EM class of security policies and security automata (Section 2.3). In Section 3 we describe our new enforcement

mechanisms – insertion automata, suppression automata and edit automata. For each mechanism, we analyze the class of security policies that the mechanism is able to enforce and provide practical examples of policies that fall in that class. In Section 4 we provide the syntax and operational semantics of a simple security policy language that admits editing operations on the instruction stream. In Section 5 we discuss some unanswered questions and our continuing research. Section 6 concludes the paper with a taxonomy of security policies.

2 Security Policies and Enforcement Mechanisms

In this section, we explain our model of software systems and how they execute, which is based on the work of Alpern and Schneider [AS87, Sch00]. We define what it means to be a security policy and give definitions for safety, liveness and EM policies. We give a new presentation of Schneider's security automata and their semantics that emphasizes our view of these machines as sequence transformers rather than property recognizers. Finally, we provide definitions of what it means for an automaton to *enforce* a property precisely and conservatively, and also what it means for one automaton to be a more effective enforcer than another automaton for a particular property.

2.1 Systems, Executions and Policies

We specify software systems at a high level of abstraction. A *system* $\mathcal{S} = (\mathcal{A}, \Sigma)$ is specified via a set of *program actions* $\mathcal{A}$ (also referred to as events or program operations) and a set of possible executions Σ. An *execution* σ is simply a finite sequence of actions $a_1, a_2, \ldots, a_n$. Previous authors have considered infinite executions as well as finite ones. We restrict ourselves to finite, but arbitrarily long executions to simplify our analysis. We use the metavariables σ and τ to range over finite sequences.

The symbol $\cdot$ denotes the empty sequence. We use the notation $\sigma[i]$ to denote the ith action in the sequence (beginning the count at 0). The notation $\sigma[..i]$ denotes the subsequence of σ involving the actions $\sigma[0]$ through $\sigma[i]$, and $\sigma[i+1..]$ denotes the subsequence of σ involving all other actions. We use the notation $\tau;\sigma$ to denote the concatenation of two sequences. When τ is a prefix of σ we write $\tau \prec \sigma$.

In this work, it will be important to distinguish between uniform systems and nonuniform systems. $(\mathcal{A}, \Sigma)$ is a *uniform* system if $\Sigma = \mathcal{A}^\star$ where $\mathcal{A}^\star$ is the set of all finite sequences of symbols from $\mathcal{A}$. Conversely, $(\mathcal{A}, \Sigma)$ is a *nonuniform* system if $\Sigma \subset \mathcal{A}^\star$. Uniform systems arise naturally when a program is completely unconstrained; unconstrained programs may execute operations in any order. However, an effective security system will often combine static program analysis and preprocessing with run-time security monitoring. Such is the case in Java virtual machines, for example, which combine type checking with stack inspection. Program analysis and preprocessing can give rise to nonuniform systems. In this paper, we are not concerned with how nonuniform

systems may be generated, be it by model checking programs, control or dataflow analysis, program instrumentation, type checking, or proof-carrying code; we care only that they exist.

A *security policy* is a predicate P on sets of executions. A set of executions Σ satisfies a policy P if and only if $P(\Sigma)$. Most common extensional program properties fall under this definition of security policy, including the following.

- *Access Control* policies specify that no execution may operate on certain resources such as files or sockets, or invoke certain system operations.
- *Availability* policies specify that if a program acquires a resource during an execution, then it must release that resource at some (arbitrary) later point in the execution.
- *Bounded Availability* policies specify that if a program acquires a resource during an execution, then it must release that resource by some fixed point later in the execution. For example, the resource must be released in at most ten steps or after some system invariant holds. We call the condition that demands release of the resource the *bound* for the policy.
- An *Information Flow* policy concerning inputs s_1 and outputs s_2 might specify that if $s_2 = f(s_1)$ in one execution (for some function f) then there must exist another execution in which $s_2 \neq f(s_1)$.

2.2 Security Properties

Alpern and Schneider [AS87] distinguish between *properties* and more general policies as follows. A security policy P is deemed to be a (computable) *property* when the policy has the following form.

$$P(\Sigma) = \forall \sigma \in \Sigma . \hat{P}(\sigma) \qquad \text{(PROPERTY)}$$

where $\hat{P}$ is a computable predicate on $\mathcal{A}^\star$.

Hence, a property is defined exclusively in terms of individual executions. A property may not specify a relationship between possible executions of the program. Information flow, for example, which can only be specified as a condition on a set of possible executions of a program, is not a property. The other example policies provided in the previous section are all security properties.

We implicitly assume that the empty sequence is contained in any property. For all the properties we are interested in it will always be okay not to run the program in question. From a technical perspective, this decision allows us to avoid repeatedly considering the empty sequence as a special case in future definitions of enforceable properties.

Given some set of actions $\mathcal{A}$, a predicate $\hat{P}$ over $\mathcal{A}^\star$ induces the security property $P(\Sigma) = \forall \sigma \in \Sigma . \hat{P}(\sigma)$. We often use the symbol $\hat{P}$ interchangeably as a predicate over execution sequences and as the induced property. Normally, the context will make clear which meaning we intend.

Safety Properties The *safety properties* are properties that specify that "nothing bad happens." We can make this definition precise as follows. $\hat{P}$ is a safety property if and only if for all $\sigma \in \Sigma$,

$$\neg \hat{P}(\sigma) \Rightarrow \forall \sigma' \in \Sigma.(\sigma \prec \sigma' \Rightarrow \neg \hat{P}(\sigma')) \qquad \text{(SAFETY)}$$

Informally, this definition states that once a bad action has taken place (thereby excluding the execution from the property) there is no extension of that execution that can remedy the situation. For example, access-control policies are safety properties since once the restricted resource has been accessed the policy is broken. There is no way to "un-access" the resource and fix the situation afterward.

Liveness Properties A *liveness property*, in contrast to a safety property, is a property in which nothing exceptionally bad can happen in any finite amount of time. Any finite sequence of actions can always be extended so that it lies within the property. Formally, $\hat{P}$ is a liveness property if and only if,

$$\forall \sigma \in \Sigma.\exists \sigma' \in \Sigma.(\sigma \prec \sigma' \wedge \hat{P}(\sigma')) \qquad \text{(LIVENESS)}$$

Availability is a liveness property. If the program has acquired a resource, we can always extend its execution so that it releases the resource in the next step.

Other Properties Surprisingly, Alpern and Schneider [AS87] show that any property can be decomposed into the conjunction of a safety property and a liveness property. Bounded availability is a property that combines safety and liveness. For example, suppose our bounded-availability policy states that every resource that is acquired must be released and must be released at most ten steps after it is acquired. This property contains an element of safety because there is a bad thing that may occur (*e.g.*, taking 11 steps without releasing the resource). It is not purely a safety property because there are sequences that are not in the property (*e.g.*, we have taken eight steps without releasing the resource) that may be extended to sequences that are in the property (*e.g.*, we release the resource on the ninth step).

2.3 EM

Recently, Schneider [Sch00] defined a new class of security properties called EM. Informally, EM is the class of properties that can be enforced by a *monitor* that runs in parallel with a *target program*. Whenever the target program wishes to execute a security-relevant operation, the monitor first checks its policy to determine whether or not that operation is allowed. If the operation is allowed, the target program continues operation, and the monitor does not change the program's behavior in any way. If the operation is not allowed, the monitor terminates execution of the program. Schneider showed that every EM property

satisfies (Safety) and hence EM is a subset of the safety properties. In addition, Schneider considered monitors for infinite sequences and he showed that such monitors can only enforce policies that obey the following continuity property.

$$\forall \sigma \in \Sigma . \neg \hat{P}(\sigma) \Rightarrow \exists i . \neg \hat{P}(\sigma[..i]) \qquad \text{(Continuity)}$$

Continuity states that any (infinite) execution that is not in the EM policy must have some finite prefix that is also not in the policy.

Security Automata Any EM policy can be enforced by a *security automaton* A, which is a deterministic finite or infinite state machine (Q, q_0, δ) that is specified with respect to some system $(\mathcal{A}, \Sigma)$. Q specifies the possible automaton states and q_0 is the initial state. The partial function $\delta : \mathcal{A} \times Q \rightarrow Q$ specifies the transition function for the automaton.

Our presentation of the operational semantics of security automata deviates from the presentation given by Alpern and Schneider because we view these machines as *sequence transformers* rather than simple *sequence recognizers*. We specify the execution of a security automaton A on a sequence of program actions σ using a labeled operational semantics.

The basic single-step judgment has the form $(\sigma, q) \xrightarrow{\tau}_A (\sigma', q')$ where σ and q denote the input program action sequence and current automaton state; σ' and q' denote the action sequence and state after the automaton processes a single input symbol; and τ denotes the sequence of actions that the automaton allows to occur (either the first action in the input sequence or, in the case that this action is "bad," no actions at all). We may also refer to the sequence τ as the observable actions or the automaton output. The input sequence σ is not considered observable to the outside world.

$\boxed{(\sigma, q) \xrightarrow{\tau}_A (\sigma', q')}$

$$(\sigma, q) \xrightarrow{a}_A (\sigma', q') \qquad \text{(A-Step)}$$

if $\sigma = a; \sigma'$
and $\delta(a, q) = q'$

$$(\sigma, q) \xrightarrow{\cdot}_A (\cdot, q) \qquad \text{(A-Stop)}$$

otherwise

We extend the single-step semantics to a multi-step semantics through the following rules.

$\boxed{(\sigma, q) \stackrel{\tau}{\Longrightarrow}_A (\sigma', q')}$

$$\frac{}{(\sigma, q) \stackrel{\cdot}{\Longrightarrow}_A (\sigma, q)} \qquad \text{(A-Reflex)}$$

$$\frac{(\sigma, q) \xrightarrow{\tau_1}_A (\sigma'', q'') \quad (\sigma'', q'') \stackrel{\tau_2}{\Longrightarrow}_A (\sigma', q')}{(\sigma, q) \stackrel{\tau_1;\tau_2}{\Longrightarrow}_A (\sigma', q')} \quad \text{(A-Trans)}$$

Limitations Erlingsson and Schneider [UES99, UES00] demonstrate that security automata can enforce important access-control policies including software fault isolation and Java stack inspection. However, they cannot enforce any of our other example policies (availability, bounded availability or information flow). Schneider [Sch00] also points out that security automata cannot enforce safety properties on systems in which the automaton cannot exert sufficient controls over the system. For example, if one of the actions in the system is the passage of time, an automaton might not be able to enforce the property because it cannot terminate an action sequence effectively — an automaton cannot stop the passage of real time.

2.4 Enforceable Properties

To be able to discuss different sorts of enforcement automata formally and to analyze how they enforce different properties, we need a formal definition of what it means for an automaton to enforce a property.

We say that an automaton A *precisely enforces* a property $\hat{P}$ on the system $(\mathcal{A}, \Sigma)$ if and only if $\forall \sigma \in \Sigma$,

1. If $\hat{P}(\sigma)$ then $\forall i.(\sigma, q_0) \stackrel{\sigma[..i]}{\Longrightarrow}_A (\sigma[i+1..], q')$ and,
2. If $(\sigma, q_0) \stackrel{\sigma'}{\Longrightarrow}_A (\cdot, q')$ then $\hat{P}(\sigma')$

Informally, if the sequence belongs to the property $\hat{P}$ then the automaton should not modify it. In this case, we say the automaton *accepts* the sequence. If the input sequence is not in the property, then the automaton may (and in fact must) edit the sequence so that the output sequence satisfies the property.

Some properties are extremely difficult to enforce precisely, so, in practice, we often enforce a stronger property that implies the weaker property in which we are interested. For example, information flow is impossible to enforce precisely using run-time monitoring as it is not even a proper property. Instead of enforcing information flow, an automaton might enforce a simpler policy such as access control. Assuming access control implies the proper information-flow policy, we say that this automaton *conservatively enforces* the information-flow policy. Formally, an automaton conservatively enforces a property $\hat{P}$ if condition 2 from above holds. Condition 1 need not hold for an automaton to conservatively enforce a property. In other words, an automaton that conservatively enforces a property may occasionally edit an action sequence that actually obeys the policy, even though such editing is unnecessary (and potentially disruptive to the benign program's execution). Of course, any such edits should

result in an action sequence that continues to obeys the policy. Henceforth, when we use the term *enforces* without qualification (precisely, conservatively) we mean enforces precisely.

We say that automaton A_1 enforces a property $\hat{P}$ *more precisely* or *more effectively* than another automaton A_2 when either

1. A_1 accepts more sequences than A_2, or
2. The two automata accept the same sequences, but the average edit distance[1] between inputs and outputs for A_1 is less than that for A_2.

3 Beyond EM

Given our novel view of security automata as sequence transformers, it is a short step to define new sorts of automata that have greater transformational capabilities. In this section, we describe insertion automata, suppression automata and their conjunction, edit automata. In each case, we characterize the properties they can enforce precisely.

3.1 Insertion Automata

An *insertion automaton* I is a finite or infinite state machine (Q, q_0, δ, γ) that is defined with respect to some system of executions $\mathcal{S} = (\mathcal{A}, \Sigma)$. Q is the set of all possible machine states and q_0 is a distinguished starting state for the machine. The partial function $\delta : \mathcal{A} \times Q \rightarrow Q$ specifies the transition function as before. The new element is a partial function γ that specifies the insertion of a number of actions into the program's action sequence. We call this the *insertion function* and it has type $\mathcal{A} \times Q \rightarrow \vec{\mathcal{A}} \times Q$. In order to maintain the determinacy of the automaton, we require that the domain of the insertion function is disjoint from the domain of the transition function.

We specify the execution of an insertion automaton as before. The single-step relation is defined below.

$$(\sigma, q) \xrightarrow{a}_I (\sigma', q') \qquad \text{(I-Step)}$$

if $\sigma = a; \sigma'$
and $\delta(a, q) = q'$

$$(\sigma, q) \xrightarrow{\tau}_I (\sigma, q') \qquad \text{(I-Ins)}$$

if $\sigma = a; \sigma'$
and $\gamma(a, q) = \tau, q'$

[1]The *edit distance* between two sequences is the minimum number of insertions, deletions or substitutions that must be applied to either of the sequences to make them equal [Gus97].

$$(\sigma, q) \overset{\cdot}{\longrightarrow}_I (\cdot, q) \quad \text{(I-Stop)}$$

otherwise

We can extend this single-step semantics to a multi-step semantics as before.

Enforceable Properties We will examine the power of insertion automata both on uniform systems and on nonuniform systems.

Theorem 1 (Uniform I-Enforcement)
*If $\mathcal{S}$ is a uniform system and insertion automaton I precisely enforces $\hat{P}$ on $\mathcal{S}$ then $\hat{P}$ obeys (*Safety*).*

Proof: Assume (anticipating a contradiction) that an insertion automaton I enforces some property $\hat{P}$ that does not satisfy (Safety). By the definition of safety, there exists a sequence τ such that $\neg\hat{P}(\tau)$ and an extension σ such that $\hat{P}(\tau;\sigma)$. Without loss of generality, consider the action of I when it has seen an input stream consisting of all but the last symbol in τ. Now, when I is confronted with the last symbol of an input sequence with prefix τ, the automaton can do one of three things (corresponding to each of the possible operational rules).

- Case (I-Step): I accepts this symbol and waits for the next. Unfortunately, the input sequence that is being processed may be exactly τ. In this case, the automaton fails to enforce $\hat{P}$ since $\neg\hat{P}(\tau)$.
- Case (I-Ins): I inserts some sequence. By taking this action the automaton gives up on enforcing the property precisely. The input sequence might be $\tau;\sigma$, an input that obeys the property, and hence the automaton unnecessarily edited the program action stream.
- Case (I-Stop): As in case (I-Ins), the automaton gives up on precise enforcement.

Hence, no matter what the automaton might try to do, it cannot enforce $\hat{P}$ precisely and we have our contradiction. ■

If we consider nonuniform systems then the insertion automaton can enforce non-safety properties. For example, reconsider the scenario in the proof above, but this time in a carefully chosen nonuniform system $\mathcal{S}'$. In $\mathcal{S}'$, the last action of every sequence is the special stop symbol, and stop appears nowhere else in $\mathcal{S}'$. Now, assuming that the sequence τ does not end in stop (and $\neg\hat{P}(\tau;\mathsf{stop})$), our insertion automaton has a safe course of action. After seeing τ, our automaton waits for the next symbol (which must exist, since we asserted the last symbol of τ is not stop). If the next symbol is stop, it inserts σ and stops, thereby enforcing the policy. On the other hand, if the program itself continues to produce σ, the automaton need do nothing.

It is normally a simple matter to instrument programs so that they conform to the nonuniform system discussed above. The instrumentation process would insert a stop event before the program exits. Moreover, to avoid the scenario in which a non-terminating program sits in a tight loop and never commits any further security-relevant actions, we could ensure that after some time period, the automaton receives a timeout signal which also acts as a stop event.

Bounded-availability properties, which are not EM properties, have the same form as the policy considered above, and as a result, an insertion automaton can enforce many bounded-availability properties on non-uniform systems. In general, the automaton monitors the program as it acquires and releases resources. Upon detecting the bound, the automaton inserts actions that release the resources in question. It also releases the resources in question if it detects termination via a stop event or timeout.

We characterize the properties that can be enforced by an insertion automaton as follows.

Theorem 2 (Nonuniform I-Enforcement)
A property $\hat{P}$ on the system $\mathcal{S} = (\mathcal{A}, \Sigma)$ can be enforced by some insertion automaton if and only if there exists a function γ_p such that for all executions $\sigma \in \mathcal{A}^\star$, if $\neg\hat{P}(\sigma)$ then

1. *$\forall \sigma' \in \Sigma . \sigma \prec \sigma' \Rightarrow \neg\hat{P}(\sigma')$, or*
2. *$\sigma \notin \Sigma$ and $\hat{P}(\sigma; \gamma_p(\sigma))$*

Proof: (If Direction) We can construct an insertion automaton that precisely enforces any of the properties $\hat{P}$ stated above. The automaton definition follows.

- States: $q \in \mathcal{A}^\star \cup \{\mathsf{end}\}$ (the sequence of actions seen so far, or end if the automaton will stop on the next step)
- Start state: $q_0 = \cdot$ (the empty sequence)
- Transition function (δ):

  Consider processing the action a.
  If our current state q is end then stop (*i.e.*, δ and γ are undefined and hence the rule (I-Stop) applies).
  Otherwise our current state q is σ and we proceed as follows.

  – If $\hat{P}(\sigma; a)$ then we emit the action a and continue in state $\sigma; a$.
  – If $\neg\hat{P}(\sigma; a)$ and $\hat{P}(\sigma)$ and $\forall \sigma' \in \Sigma . \sigma; a \prec \sigma' \Rightarrow \neg\hat{P}(\sigma')$ then we simply stop.
  – If $\neg\hat{P}(\sigma; a)$ and $\sigma; a \notin \Sigma$ and $\hat{P}(\sigma; a; \gamma_p(\sigma; a))$. then emit a and continue in state $\sigma; a$

- Insertion function (γ): Consider processing the action a in state $q = \sigma$.

  – If $\neg\hat{P}(\sigma; a)$ and $\neg\hat{P}(\sigma)$ and $\forall \sigma' \in \Sigma . \sigma; a \prec \sigma' \Rightarrow \neg\hat{P}(\sigma')$ then insert $\gamma_p(\sigma)$ and continue in state end.

If σ is the input so far, the automaton maintains the following invariant $\textsc{Inv}_p(q)$.

- If $q = \mathsf{end}$ then the automaton has emitted $\sigma; \gamma_p(\sigma)$ and $\hat{P}(\sigma; \gamma_p(\sigma))$ and the next action is a and $\forall \sigma' \in \Sigma. \sigma; a \prec \sigma' \Rightarrow \neg\hat{P}(\sigma')$.
- Otherwise, $q = \sigma$ and σ has been emitted and either $\hat{P}(\sigma)$ or $(\neg\hat{P}(\sigma)$ and $\sigma \notin \Sigma$ and $\hat{P}(\sigma; \gamma_p(\sigma)))$.

The automaton can initially establish $\textsc{Inv}_p(q_0)$ since our definition of a property assumes $\hat{P}(\cdot)$ for all properties. A simple inductive argument on the length of the input σ suffices to show that the invariant is maintained for all inputs.

Given this invariant, it is straightforward to show that the automaton processes every input $\sigma \in \Sigma$ properly and precisely enforces $\hat{P}$. There are two cases.

- **Case** $\hat{P}(\sigma)$:
  Consider any prefix $\sigma[..i]$. By induction on i, we show the automaton accepts $\sigma[..i]$ without stopping, inserting any actions or moving to the state end.
  If $\hat{P}(\sigma[..i])$ then the automaton accepts this prefix and continues.
  If $\neg\hat{P}(\sigma[..i])$ then since $\sigma[..i] \prec \sigma$ (and $\hat{P}(\sigma)$), it must be the case that $\sigma[..i] \notin \Sigma$ and $\hat{P}(\sigma[..i]; \gamma_p(\sigma[..i]))$. Hence, the automaton accepts this prefix and continues.
- **Case** $\neg\hat{P}(\sigma)$:
  $\textsc{Inv}_p$ and the automaton definition imply that whenever the automaton halts (because of lack of input or because it stops intentionally), $\hat{P}(\sigma_o)$ where σ_o is the sequence of symbols that have been output. Hence, the automaton processes this input properly as well.

(Only-If Direction) Define $\gamma_p(\sigma)$ to be some τ such that $\hat{P}(\sigma; \tau)$ (and undefined if no such τ exists). We consider any arbitrary $\sigma \in \mathcal{A}^\star$ such that $\neg\hat{P}(\sigma)$ and show that one of the following must hold.

1. $\forall \sigma' \in \Sigma. \sigma \prec \sigma' \Rightarrow \neg\hat{P}(\sigma')$, or
2. $\sigma \notin \Sigma$ and $\hat{P}(\sigma; \gamma_p(\sigma))$

Consider any case in which 1. does not hold. Then, $\exists \sigma' \in \Sigma. \sigma \prec \sigma' \wedge \hat{P}(\sigma')$, so by definition of γ_p, $\hat{P}(\sigma; \gamma_p(\sigma))$. All that remains to show that 2. holds in any case in which 1. does not is that $\sigma \notin \Sigma$ (given that some insertion automaton I enforces $\hat{P}$, $\neg\hat{P}(\sigma)$, $\sigma \prec \sigma'$, and $\hat{P}(\sigma')$).

By the first part of the definition of precise enforcement, I must not make any edits when supplied σ' as input. This excludes the use of rules I-Ins and I-Stop, so I only steps via rule I-Step on input σ'. At any point during I's processing of σ (where $\sigma \prec \sigma'$), exactly these same I-Step transitions must be made because the input sequence may later be extended to σ'. Therefore,

$(\sigma, q_0) \stackrel{\sigma}{\Longrightarrow}_I (\cdot, q')$ for some q'. Assume (for the sake of obtaining a contradiction) that $\sigma \in \Sigma$. Then, by the second part of the definition of precise enforcement and the fact that $(\sigma, q_0) \stackrel{\sigma}{\Longrightarrow}_I (\cdot, q')$, we have $\hat{P}(\sigma)$, which violates the assumption that $\neg\hat{P}(\sigma)$. Thus, $\sigma \notin \Sigma$ as required, and case 2. above must hold whenever 1. does not. ■

Limitations Like the security automaton, the insertion automaton is limited by the fact that it may not be able to be able to exert sufficient controls over a system. More precisely, it may not be possible for the automaton to synthesize certain events and inject them into the action stream. For example, an automaton may not have access to a principal's private key. As a result, the automaton may have difficulty enforcing a fair exchange policy that requires two computational agents to exchange cryptographically signed documents. Upon receiving a signed document from one agent, the insertion automaton may not be able to force the other agent to sign the second document and it cannot forge the private key to perform the necessary cryptographic operations itself.

3.2 Suppression Automata

A *suppression automaton* S is a state machine (Q, q_0, δ, ω) that is defined with respect to some system of executions $\mathcal{S} = (\mathcal{A}, \Sigma)$. As before, Q is the set of all possible machine states, q_0 is a distinguished starting state for the machine and the partial function δ specifies the transition function. The partial function $\omega : \mathcal{A} \times Q \rightarrow \{-, +\}$ has the same domain as δ and indicates whether or not the action in question is to be suppressed $(-)$ or emitted $(+)$.

$$(\sigma, q) \stackrel{a}{\longrightarrow}_S (\sigma', q') \qquad \text{(S-StepA)}$$

if $\sigma = a; \sigma'$
and $\delta(a, q) = q'$
and $\omega(a, q) = +$

$$(\sigma, q) \stackrel{\cdot}{\longrightarrow}_S (\sigma', q') \qquad \text{(S-StepS)}$$

if $\sigma = a; \sigma'$
and $\delta(a, q) = q'$
and $\omega(a, q) = -$

$$(\sigma, q) \stackrel{\cdot}{\longrightarrow}_S (\cdot, q) \qquad \text{(S-Stop)}$$

otherwise

We extend the single-step relation to a multi-step relation using the reflexivity and transitivity rules from above.

Enforceable Properties In a uniform system, suppression automata can only enforce safety properties.

Theorem 3 (Uniform S-Enforcement)
*If $\mathcal{S}$ is a uniform system and suppression automaton S precisely enforces $\hat{P}$ on $\mathcal{S}$ then $\hat{P}$ obeys (*SAFETY*).*

Proof (sketch): The argument is similar to the argument for insertion automata given in the previous section. If we are attempting to enforce a property $\hat{P}$, we cannot allow any sequence τ such that $\neg\hat{P}(\tau)$, even though there may be an extension σ such that $\hat{P}(\tau;\sigma)$. Any step of S when processing the final symbol of τ would result in either accepting τ (despite the fact that $\neg\hat{P}(\tau)$) or giving up on precise enforcement altogether. ■

In a nonuniform system, suppression automata can once again enforce non-EM properties. For example, consider the following system $\mathcal{S}$.

$$\begin{array}{rcl} \mathcal{A} & = & \{\mathsf{aq}, \mathsf{use}, \mathsf{rel}\} \\ \Sigma & = & \{\mathsf{aq}; \mathsf{rel}, \\ & & \ \mathsf{aq}; \mathsf{use}; \mathsf{rel}, \\ & & \ \mathsf{aq}; \mathsf{use}; \mathsf{use}; \mathsf{rel}\} \end{array}$$

The symbols $\mathsf{aq}, \mathsf{use}, \mathsf{rel}$ denote acquisition, use and release of a resource. The set of executions includes zero, one, or two uses of the resource. Such a scenario might arise were we to publish a policy that programs can use the resource at most two times. After publishing such a policy, we might find a bug in our implementation that makes it impossible for us to handle the load we were predicting. Naturally we would want to tighten the security policy as soon as possible, but we might not be able to change the policy we have published. Fortunately, we can use a suppression automaton to suppress extra uses and dynamically change the policy from a two-use policy to a one-use policy. Notice that an ordinary security automaton is not sufficient to make this change because it can only terminate execution.[2] After terminating a two-use application, it would be unable to insert the release necessary to satisfy the policy.

We can also compare the power of suppression automata with insertion automata. A suppression automaton cannot enforce the bounded-availability policy described in the previous section because it cannot insert release events that are necessary if the program halts prematurely. That is, although the suppression automaton could suppress all non-release actions upon reaching the bound (waiting for the release action to appear), the program may halt without releasing, leaving the resource unreleased. Note also that the suppression automaton

[2]Premature termination of these executions takes us outside the system $\mathcal{S}$ since the rel symbol would be missing from the end of the sequence. To model the operation of a security automaton in such a situation we would need to separate the set of possible input sequences from the set of possible output sequences. For the sake of simplicity, we have not done so in this paper.

cannot simply suppress resource acquisitions and uses because this would modify sequences that actually do satisfy the policy, contrary to the definition of precise enforcement. Hence, insertion automata can enforce some properties that suppression automata cannot.

For any suppression automaton, we can construct an insertion automaton that enforces the same property. The construction proceeds as follows. While the suppression automaton acts as a simple security automaton, the insertion automaton can clearly simulate it. When the suppression automaton decides to suppress an action a, it does so because there exists some extension σ of the input already processed (τ) such that $\hat{P}(\tau;\sigma)$ but $\neg\hat{P}(\tau;a;\sigma)$. Hence, when the suppression automaton suppresses a (giving up on precisely enforcing any sequence with $\sigma;a$ as a prefix), the insertion automaton merely inserts σ and terminates (also giving up on precise enforcement of sequences with $\sigma;a$ as a prefix). Of course, in practice, if σ is uncomputable or only intractably computable from τ, suppression automata are useful.

There are also many scenarios in which suppression automata are *more precise* enforcers than insertion automata. In particular, in situations such as the one described above in which we publish one policy but later need to restrict it due to changing system requirements or policy bugs, we can use suppression automata to suppress resource requests that are no longer allowed. Each suppression results in a new program action stream with an edit distance increased by 1, whereas the insertion automaton may produce an output with an arbitrary edit distance from the input.

Before we can characterize the properties that can be enforced by a suppression automaton, we must generalize our suppression functions so they act over sequences of symbols. Given a set of actions $\mathcal{A}$, a computable function $\omega^\star : \mathcal{A}^\star \to \mathcal{A}^\star$ is a *suppression function* if it satisfies the following conditions.

1. $\omega^\star(\cdot) = \cdot$
2. $\omega^\star(\sigma;a) = \omega^\star(\sigma);a$, or
   $\omega^\star(\sigma;a) = \omega^\star(\sigma)$

A suppression automaton can enforce the following properties.

Theorem 4 (Nonuniform S-Enforcement)
A property $\hat{P}$ on the system $\mathcal{S} = (\mathcal{A}, \Sigma)$ is enforceable by a suppression automaton if and only if there exists a suppression function $\omega^\star$ such that for all sequences $\sigma \in \mathcal{A}^\star$,

- *if $\hat{P}(\sigma)$ then $\omega^\star(\sigma) = \sigma$, and*
- *if $\neg\hat{P}(\sigma)$ then*
  1. *$\forall\sigma' \in \Sigma.\sigma \prec \sigma' \Rightarrow \neg\hat{P}(\sigma')$, or*
  2. *$\sigma \notin \Sigma$ and $\forall\sigma' \in \Sigma.\sigma \prec \sigma' \Rightarrow \hat{P}(\omega^\star(\sigma'))$*

Proof: (If Direction) As in the previous section, given one of the properties $\hat{P}$ described above, we can construct a suppression automaton that enforces it.

- States: $q \in \mathcal{A}^\star \times \{+,-\}$ (the sequence of actions seen so far paired with $+$ $(-)$ to indicate that no actions have (at least one action has) been suppressed so far)
- Start state: $q_0 = \langle \cdot, + \rangle$
- Transition function (for simplicity, we combine δ and ω):

  Consider processing the action a.
  If the current state q is $\langle \sigma, + \rangle$ then

  - If $\hat{P}(\sigma; a)$, then we emit the action a and continue in state $\langle \sigma; a, + \rangle$.
  - If $\neg\hat{P}(\sigma; a)$ and $\forall \sigma' \in \Sigma. \sigma; a \prec \sigma' \Rightarrow \neg\hat{P}(\sigma')$ then we simply halt.
  - Otherwise (*i.e.*, $\neg\hat{P}(\sigma; a)$ and $\exists \sigma' \in \Sigma. \sigma; a \prec \sigma' \wedge \hat{P}(\sigma')$),
    - if $\omega^\star(\sigma; a) = \omega^\star(\sigma)$ we suppress a and continue in state $\langle \sigma; a, - \rangle$.
    - and finally, if $\omega^\star(\sigma; a) = \omega^\star(\sigma); a$ we emit a and continue in state $\langle \sigma; a, + \rangle$.

  Otherwise our current state q is $\langle \sigma, - \rangle$.

  - If $\hat{P}(\omega^\star(\sigma))$ then stop.
  - Otherwise (*i.e.*, $\neg\hat{P}(\omega^\star(\sigma))$),
    - if $\omega^\star(\sigma; a) = \omega^\star(\sigma)$ then suppress a and continue in state $\langle \sigma; a, - \rangle$.
    - and finally, if $\omega^\star(\sigma; a) = \omega^\star(\sigma); a$ then emit a and continue in state $\langle \sigma; a, - \rangle$.

If σ is the input so far, the automaton maintains the following invariant $\textsc{Inv}_p(q)$.

- If $q = \langle \sigma, + \rangle$ then
  1. $\omega^\star(\sigma)$ has been emitted.
  2. $(\hat{P}(\sigma)$ and $\omega^\star(\sigma) = \sigma)$ or $(\neg\hat{P}(\sigma)$ and $\omega^\star(\sigma) = \sigma$ and $\sigma \notin \Sigma$ and $\forall \sigma' \in \Sigma. \sigma \prec \sigma' \Rightarrow \hat{P}(\omega^\star(\sigma')))$.
- If $q = \langle \sigma, - \rangle$ then
  1. $\omega^\star(\sigma)$ has been emitted.
  2. $\hat{P}(\omega^\star(\sigma))$ or $(\neg\hat{P}(\omega^\star(\sigma))$ and $\forall \sigma' \in \Sigma. \sigma \prec \sigma' \Rightarrow \hat{P}(\omega^\star(\sigma'))$ and $\sigma \notin \Sigma)$.

The automaton can initially establish $\textsc{Inv}_p(q_0)$ since our definition of a property assumes $\hat{P}(\cdot)$ for all properties and $\omega^\star(\cdot) = \cdot$ for all suppression functions. A simple inductive argument on the length of the input σ suffices to show that the invariant is maintained for all inputs.

Given this invariant, it is straightforward to show that the automaton processes every input $\sigma \in \Sigma$ properly and precisely enforces $\hat{P}$. There are two cases.

- **Case** $\hat{P}(\sigma)$:
  This case is similar to the analogous case for insertion automata. We prove the automaton accepts the input without stopping or suppressing any actions by induction on the length of the sequence.

- **Case** $\neg\hat{P}(\sigma)$:
  As before, $\text{I}\textsc{nv}_p$ implies the automaton always stops in the state in which the automaton output σ_o satisfies the property. This implies we process σ properly.

(Only-If Direction) Define $\omega^\star(\sigma)$ to be whatever sequence is emitted by the suppression automaton S on input σ. We first show that this is indeed a suppression function. Clearly, $\omega^\star(\cdot) = \cdot$. When processing some action a after having already processed any sequence σ, the automaton may step via S-StepA, S-StepS, or S-Stop. In the S-StepA case, the automaton emits whatever has been emitted in processing σ (by definition, this is $\omega^\star(\sigma)$), followed by a. In the other cases, the automaton emits only $\omega^\star(\sigma)$. Hence, $\omega^\star$ is a valid suppression function.

Now consider any arbitrary $\sigma \in \mathcal{A}^\star$. By the definition of precise enforcement, automaton S does not modify any sequence σ such that $\hat{P}(\sigma)$, so we have satisfied the requirement that if $\hat{P}(\sigma)$ then $\omega^\star(\sigma) = \sigma$. All that remains is to show that if $\neg\hat{P}(\sigma)$ then

1. $\forall \sigma' \in \Sigma. \sigma \prec \sigma' \Rightarrow \neg\hat{P}(\sigma')$, or
2. $\sigma \notin \Sigma$ and $\forall \sigma' \in \Sigma. \sigma \prec \sigma' \Rightarrow \hat{P}(\omega^\star(\sigma'))$

Consider any case where $\neg\hat{P}(\sigma)$ and 1. does not hold (*i.e.*, $\exists \sigma' \in \Sigma. \sigma \prec \sigma' \wedge \hat{P}(\sigma')$). Analogous to the case for insertion automata, because $\hat{P}(\sigma')$, S may only process σ' with S-StepA and therefore only processes σ (which may later be extended to σ') with S-StepA. Hence, $(\sigma, q_0) \stackrel{\sigma}{\Longrightarrow}_S (\cdot, q')$ for some q'. Since S enforces $\hat{P}$ and $\neg\hat{P}(\sigma)$, it must be the case that $\sigma \notin \Sigma$.

Finally, by the definition of $\omega^\star$ and the second part of the definition of precise enforcement, $\forall \sigma' \in \Sigma. \hat{P}(\omega^\star(\sigma'))$, implying in case 2. above that $\forall \sigma' \in \Sigma. \sigma \prec \sigma' \Rightarrow \hat{P}(\omega^\star(\sigma'))$. Therefore, 2. indeed holds whenever 1. does not. ∎

Limitations Similarly to its relatives, a suppression automaton is limited by the fact that some events may not be suppressible. For example, the program may have a direct connection to some output device and the automaton may be unable to interpose itself between the device and the program. It might also be the case that the program is unable to continue proper execution if an action is suppressed. For instance, the action in question might be an input operation.

3.3 Edit Automata

We form an *edit automaton* E by combining the insertion automaton with the suppression automaton. Our machine is now described by a 5-tuple with the form $(Q, q_0, \delta, \gamma, \omega)$. The operational semantics are derived from the composition of the operational rules from the two previous automata.

$$(\sigma, q) \xrightarrow{a}_E (\sigma', q') \qquad \text{(E-StepA)}$$

if $\sigma = a; \sigma'$
and $\delta(a, q) = q'$
and $\omega(a, q) = +$

$$(\sigma, q) \xrightarrow{\cdot}_E (\sigma', q') \qquad \text{(E-StepS)}$$

if $\sigma = a; \sigma'$
and $\delta(a, q) = q'$
and $\omega(a, q) = -$

$$(\sigma, q) \xrightarrow{\tau}_E (\sigma, q') \qquad \text{(E-Ins)}$$

if $\sigma = a; \sigma'$
and $\gamma(a, q) = \tau, q'$

$$(\sigma, q) \xrightarrow{\cdot}_E (\cdot, q) \qquad \text{(E-Stop)}$$

otherwise

We again extend this single-step semantics to a multi-step semantics with the rules for reflexivity and transitivity.

Enforceable Properties As with insertion and suppression automata, edit automata are only capable of enforcing safety properties in uniform systems.

Theorem 5 (Uniform E-Enforcement)
*If $\mathcal{S}$ is a uniform system and edit automaton E precisely enforces $\hat{P}$ on $\mathcal{S}$ then $\hat{P}$ obeys (*Safety*).*

Proof (sketch): The argument is similar to that given in the proofs of Uniform I- and S-Enforcement. When confronted with processing the final action of τ such that $\neg\hat{P}(\tau)$ but $\hat{P}(\tau; \sigma)$ for some σ, E must either accept τ (despite the fact that $\neg\hat{P}(\tau)$) or give up on precise enforcement altogether. ■

The following theorem provides the formal basis for the intuition given above that insertion automata are strictly more powerful than suppression automata. Because insertion automata enforce a superset of properties enforceable by suppression automata, edit automata (which are a composition of insertion and suppression automata) precisely enforce exactly those properties that are precisely enforceable by insertion automata.

Theorem 6 (Nonuniform E-Enforcement)
A property $\hat{P}$ on the system $\mathcal{S} = (\mathcal{A}, \Sigma)$ can be enforced by some edit automaton if and only if there exists a function γ_p such that for all executions $\sigma \in \mathcal{A}^\star$, if $\neg\hat{P}(\sigma)$ then

1. *$\forall \sigma' \in \Sigma . \sigma \prec \sigma' \Rightarrow \neg\hat{P}(\sigma')$, or*
2. *$\sigma \notin \Sigma$ and $\hat{P}(\sigma; \gamma_p(\sigma))$*

Proof: (If Direction) By the Nonuniform I-Enforcement theorem, given any function γ_p and property $\hat{P}$ satisfying the requirements stated above, we may build an insertion automaton $I = (Q, q_0, \delta, \gamma)$ to enforce $\hat{P}$. Then, we can construct an edit automaton $E = (Q, q_0, \delta, \gamma, \omega)$, where ω is defined to be $+$ over all of its domain (the same domain as δ). Clearly, E and I enforce the same property because whenever I steps via I-Step, I-Ins, or I-Stop, E respectively steps via E-StepA, E-Ins, or E-Stop, while emitting exactly the same actions as I. Because I enforces $\hat{P}$, so too must E.

(Only-If Direction) This direction proceeds exactly as the Only-If Direction of the proof of Nonuniform I-Enforcement. ■

Although edit automata are no more powerful precise enforcers than insertion automata, we can very effectively enforce a wide variety of security policies conservatively with edit automata. We describe a particularly important application, the implementation of transactions policies, in the following section.

3.4 An Example: Transactions

To demonstrate the power of our edit automata, we show how to implement the monitoring of transactions. The desired properties of atomic transactions [EN94], commonly referred to as the ACID properties, are atomicity (either the entire transaction is executed or no part of it is executed), consistency preservation (upon completion of the transaction the system must be in a consistent state), isolation (the effects of a transaction should not be visible to other concurrently executing transactions until the first transaction is committed), and durability or permanence (the effects of a committed transaction cannot be undone by a future failed transaction).

The first property, atomicity, can be modeled using an edit automaton by suppressing input actions from the start of the transaction. If the transaction completes successfully, the entire sequence of actions is emitted atomically to the output stream; otherwise it is discarded. Consistency preservation can be enforced by simply verifying that the sequence to be emitted leaves the system in a consistent state. The durability or permanence of a committed transaction is ensured by the fact that committing a transaction is modeled by outputting the corresponding sequence of actions to the output stream. Once an action has been written to the output stream it can no longer be touched by the automaton;

furthermore, failed transactions output nothing. We only model the actions of a single agent in this example and therefore ignore issues of isolation.

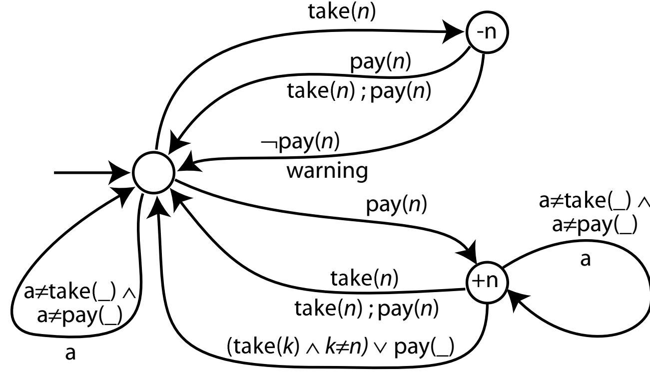

Figure 1: An edit automaton to enforce the market policy conservatively.

To make our example more concrete, we will model a simple market system with two main actions, take(n) and pay(n), which represent acquisition of n apples and the corresponding payment. We let `a` range over other actions that might occur in the system (such as `window-shop` or `browse`). Our policy is that every time an agent takes n apples it must pay for those apples. Payments may come before acquisition or vice versa. The automaton conservatively enforces atomicity of this transaction by emitting take(n); pay(n) only when the transaction completes. If payment is made first, the automaton allows clients to perform other actions such as `browse` before paying (the pay-take transaction appears atomically after all such intermediary actions). On the other hand, if apples are taken and not paid for immediately, we issue a warning and abort the transaction. Consistency is ensured by remembering the number of apples taken or the size of the prepayment in the state of the machine. Once acquisition and payment occur, the sale is final and there are no refunds (durability).

Figure 1 displays the edit automaton that conservatively enforces our market policy. The nodes in the picture represent the automaton states and the arcs represent the transitions. When a predicate above an arc is true, the transition will be taken. The sequence below an arc represents the actions that are emitted. Hence, an arc with no symbols below it is a suppression transition. An arc with multiple symbols below it is an insertion transition.

4 A Simple Policy Language

In this section we describe the syntax and semantics of a simple language for specifying edit automata.

4.1 Syntax

A program in our language consists of the declaration of automaton state and transition function. The state is described by some number of state variables $s_1, \ldots, s_n$, which are assigned some initial values $v_1, \ldots, v_n$. We leave the value space unspecified. Normally, it would include all values found in a full-fledged programming language.

We specify two transition functions in order to differentiate between *internal* events, which are events or program actions under the control of the security automaton, and *external* events, which the security automaton cannot suppress (or synthesize and insert). We use the metavariable i_{in} to denote internal events and i_{ex} to denote external events. Each transition function is a sequence of nested if-then-else statements. The guard of each statement consists of boolean predicates over the current state and the most recently read instruction. Guards can be combined using standard boolean algebra. We assume some set of atomic predicates p and a suitable evaluation function for these predicates. If, when evaluating a transition function, none of the guards evaluate to true then the automaton halts.

The main body of the transition function consists of assignments to state variables and commands to emit a sequence of instructions. We assume some set of functions f that compute values to be stored in a state variable or instructions to be emitted by the automaton. Each sequence of commands ends with a command to continue evaluation in the next state without consuming the current input symbol (**next**), to continue in the next state and to consume the input (**consume;next**), or to halt (**halt**).

The following BNF grammar describes the language syntax.

$$
\begin{array}{lcl}
P & ::= & \text{let} \\
 & & \quad \text{states}: \\
 & & \qquad s_1 = v_1 \\
 & & \qquad \vdots \\
 & & \qquad s_n = v_n \\
 & & \quad \text{transitions}: \\
 & & \qquad \lambda i_{ex}.T \\
 & & \qquad \lambda i_{in}.T \\
 & & \text{in} \\
 & & \quad \textbf{run} \\
 & & \text{end}
\end{array}
\qquad
\begin{array}{lcl}
T & ::= & \text{if } B \text{ then } S \text{ else } T \mid \textbf{halt} \\
 & & \\
B & ::= & p(s_1, \ldots, s_n, i) \\
 & \mid & B_1 \wedge B_2 \mid B_1 \vee B_2 \mid \neg B \\
 & & \\
S & ::= & s_k \ = \ G \ ; S \\
 & \mid & \textbf{emit}(G) \ ; S \\
 & \mid & \textbf{next} \\
 & \mid & \textbf{consume}; \textbf{next} \\
 & \mid & \textbf{halt} \\
 & & \\
G & ::= & f(s_1, \ldots, s_n, i)
\end{array}
$$

4.2 Operational Semantics

In order to specify the execution of our program, we need to define the *system configurations* that arise during computation. A system configuration M is a 4-tuple that contains the instruction stream being read by the automaton, a

$$\boxed{(\sigma, q, F, C) \overset{\tau}{\longmapsto} (\sigma', q', F, C')}$$

$$\frac{a \in \mathit{external}}{(a;\sigma, q, (\lambda j_{ex}.t, \lambda j_{in}.t'), \cdot) \overset{\cdot}{\longmapsto} (a;\sigma, q, (\lambda j_{ex}.t, \lambda j_{in}.t'), [a/j_{ex}]t)} \quad \text{(Sem-Lam-Ex)}$$

$$\frac{a \in \mathit{internal}}{(a;\sigma, q, (\lambda j_{ex}.t, \lambda j_{in}.t'), \cdot) \overset{\cdot}{\longmapsto} (a;\sigma, q, (\lambda j_{ex}.t, \lambda j_{in}.t'), [a/j_{in}]t')} \quad \text{(Sem-Lam-In)}$$

$$\frac{[v_1/s_1, \ldots, v_n/s_n]B \Downarrow \mathit{true}}{(\sigma, \{s_1 = v_1, \ldots, s_n = v_n\}, F, \text{if } B \text{ then } S \text{ else } T) \overset{\cdot}{\longmapsto} (\sigma, \{s_1 = v_1, \ldots, s_n = v_n\}, F, S)} \quad \text{(Sem-Bool-T)}$$

$$\frac{[v_1/s_1, \ldots, v_n/s_n]B \Downarrow \mathit{false}}{(\sigma, \{s_1 = v_1, \ldots, s_n = v_n\}, F, \text{if } B \text{ then } S \text{ else } T) \overset{\cdot}{\longmapsto} (\sigma, \{s_1 = v_1, \ldots, s_n = v_n\}, F, T)} \quad \text{(Sem-Bool-F)}$$

$$\frac{[v_1/s_1, \ldots, v_n/s_n]G \Downarrow v}{(\sigma, \{s_1 = v_1, \ldots, s_n = v_n\}, F, s_k = G \;;\; S) \overset{\cdot}{\longmapsto} (\sigma, \{s_1 = v_1, \ldots, s_n = v_n\}[s_k \mapsto v], F, S)} \quad \text{(Sem-State-Update)}$$

$$\frac{[v_1/s_1, \ldots, v_n/s_n]G \Downarrow a_1, \ldots, a_p \qquad \forall k, 1 \le k \le p, a_k \in \mathit{internal}}{(\sigma, q, F, \mathbf{emit}(G) \;;\; S) \overset{a_1, \ldots, a_p}{\longmapsto} (\sigma, q, F, S)} \quad \text{(Sem-Emit)}$$

$$\frac{a \in \mathit{external}}{(a;\sigma, q, F, \mathbf{consume}; \mathbf{next}) \overset{a}{\longmapsto} (\sigma, q, F, \cdot)} \quad \text{(Sem-Consume-Ex)}$$

$$\frac{a \in \mathit{internal}}{(a;\sigma, q, F, \mathbf{consume}; \mathbf{next}) \overset{\cdot}{\longmapsto} (\sigma, q, F, \cdot)} \quad \text{(Sem-Consume-In)}$$

$$\frac{}{(\sigma, q, F, \mathbf{next}) \overset{\cdot}{\longmapsto} (\sigma, q, F, \cdot)} \quad \text{(Sem-Next)}$$

$$\frac{a \in \mathit{external}}{(a;\sigma, q, F, \mathbf{halt}) \overset{a}{\longmapsto} (\mathit{nil}, q, F, \cdot)} \quad \text{(Sem-Halt-Ex)}$$

$$\frac{a \in \mathit{internal}}{(a;\sigma, q, F, \mathbf{halt}) \overset{\cdot}{\longmapsto} (\mathit{nil}, q, F, \cdot)} \quad \text{(Sem-Halt-In)}$$

$$\boxed{(\sigma, q, F, C) \overset{\tau}{\longmapsto}{}^{*} (\sigma', q', F, C')}$$

$$\frac{}{(\sigma, q, F, C) \overset{\cdot}{\longmapsto}{}^{*} (\sigma, q, F, C)} \quad \text{(Sem-Refl)}$$

$$\frac{(\sigma, q, F, C) \overset{\tau_1}{\longmapsto} (\sigma', q', F, C') \qquad (\sigma', q', F, C') \overset{\tau_2}{\longmapsto}{}^{*} (\sigma'', q'', F, C'')}{(\sigma, q, F, C) \overset{\tau_1;\tau_2}{\longmapsto}{}^{*} (\sigma'', q'', F, C'')} \quad \text{(Sem-Trans)}$$

$$\boxed{(\sigma, P) \overset{\tau}{\Longrightarrow} (\sigma', q, F, C)}$$

$$\frac{(\sigma, \{s_1 = v_1, \ldots, s_n = v_n\}, (\lambda j_{ex}.t,\ \lambda j_{in}.t'), \cdot\ ) \overset{\tau}{\longmapsto}{}^{*} (\sigma', q, F, C)}{(\sigma, P) \overset{\tau}{\Longrightarrow} (\sigma', q, F, C)} \quad \text{(Sem-Setup)}$$

where $P = \text{let states} : s_1 = v_1, \ldots, s_n = v_n \;;\; \text{transitions} : \lambda j_{ex}.t,\ \lambda j_{in}.t' \text{ in } \mathbf{run} \text{ end}$

Figure 2: Operational semantics.

set of state variables, the pair of functions that describe transitions for input symbols, and the program expression currently being evaluated. We use the metavariable F to range over the pair of functions and the metavariable C to range over the program expression under evaluation.

When paired with an action stream, the program gives rise to a machine configuration in natural way. We specify the execution of our language using the following judgments.

$M_1 \overset{\tau}{\longmapsto} M_2$	configuration M_1 emits τ when evaluating to configuration M_2
$M_1 \overset{\tau}{\longmapsto}^* M_2$	configuration M_1 emits τ when evaluating to configuration M_2 (in many steps)
$(\sigma, P) \overset{\tau}{\Longrightarrow} M$	sequence σ and program P emit τ when evaluating to configuration M (in many steps)
$B \Downarrow b$	B evaluates to boolean b
$G \Downarrow v$	G evaluates to v

The operational semantics of our language are given in Figure 2.

Rules Sem-Lam-Ex and Sem-Lam-In describe how the automaton begins processing the topmost instruction of the input stream. Control flow differs based on whether the instruction is external or internal. We begin evaluating the appropriate transition expression by substituting the current instruction for the formal parameter that represents it. We postpone removing the instruction from the instruction stream because we will later wish to use it to guide evaluation according to the type of the instruction.

If the program expression currently being evaluated is a guarded statement, we begin by reducing the boolean guard to true or false as described by rules Sem-Bool-T and Sem-Bool-F. Since the boolean guard contains predicates over the the current state, this involves substituting into it the current values of the state variables. Evaluating the guards may also involve some standard boolean algebra. Due to space constraints, we have omitted these standard rules.

If a guard evaluates to true, we evaluate the body of the guarded statement. This may involve updates of state variables by functions parameterized over the current state. Rule Sem-State-Update describes the update of a state variable and the necessary substitutions. In addition, the statement may include emitting instructions to the output stream. Rule Sem-Emit describes the emission of internal instructions. The emission of external instructions is not under the control of the programmer and is done automatically as described by rules Sem-Consume-Ex and Sem-Halt-Ex.

The last step in evaluating the transition function can be to consume the instruction from the input stream before starting to process the next instruction, to reevaluate the transition function with the current instruction and the updated state, or to halt the automaton. As mentioned above, if the instruction read from the input stream is external, it is automatically emitted to the output

stream before it is consumed. In order to halt the automaton, we transition to a system configuration in which the rest of the input stream is ignored.

4.3 Example

As an illustration of the use of our language to encode a security policy, we show a possible way of encoding the transaction automaton. The code in Figure 3 corresponds exactly to the picture of the automaton in Figure 1.

```
let
 states: apples = 0
 transitions:
   λ i_ex.halt
   λ i_in.
     if ValidTransition(apples, i_in)
     then apples =
            if(apples=0 ∧ i_in=take(n)) then -n
            else if (apples=0
                     ∧ i_in=pay(n)) then +n
            else if (apples=+n ∧ i_in ≠ take(_)
                     ∧ i_in ≠ pay(_)) then +n
            else 0;
          emit(getEmitActions(apples, i_in));
          consume;
          next
      else halt
in run end


getEmitActions(apples, i_in) =
  if (apples ≥ 0 ∧ i_in ≠ take(_)
      ∧ i_in ≠ pay(_)) then i_in
  else if (apples = -n ∧ i_in=pay(n)) then
             take(n);pay(n)
  else if (apples = -n ∧ i_in ≠pay(n)) then
             warning
  else if (apples = +n ∧ i_in=take(n)) then
             take(n);pay(n)
  else ·
```

Figure 3: Specification of transaction automaton in Figure 1.

We assume the predicate `ValidTransition(`q`,`a`)` evaluates to true if and only if a transition on input action a in state q is defined in the market automaton. On valid transitions the state variable *apples* is updated, after which the appropriate instructions, as calculated by function *getEmitActions*, are emitted to the output stream.

4.4 Adequacy of Policy Language

Our policy language is sufficiently expressive to describe any edit automaton. Consider the edit automaton $E = (Q, q_0, \delta, \gamma, \omega)$. We can encode this automaton

```
let
  states : s = q_0
  transitions : λi_ex. halt
                λi_in. T
in
  run
end
```

where $T =$

```
if (dom_δ (i_in, s) ∧ (ω(i_in, s) = +))
then s = δ(i_in, s); emit(i_in); consume; next
if (dom_δ (i_in, s) ∧ (ω(i_in, s) = −))
then s = δ(i_in, s); consume; next
if (dom_γ (i_in, s))
then
  s = γ_state(i_in, s);
  emit(γ_seq(i_in, s));
  next
else halt
```

Figure 4: General encoding of an edit automaton.

in the policy language provided we assume the existence of a set of values for encoding automaton states as well as the following auxiliary functions and predicates.

$\mathsf{dom}_\delta(a, q)$	true if and only if δ is defined on a, q
$\mathsf{dom}_\gamma(a, q)$	true if and only if γ is defined on a, q
$\delta(a, q)$	implements E's transition function
$\omega(a, q)$	implements E's suppression function
$\gamma_{state}(a, q)$	produces the next state of γ
$\gamma_{seq}(a, q)$	produces the inserted actions of γ

The encoding (Figure 4) uses a single variable s to represent the current state. We have been implicitly assuming that all actions processed by our edit automata are internal actions susceptible to insertion and suppression. Hence, our program's external transition function is trivial. The internal transition function encodes the operations of the automaton rather directly.

5 Future Work

We are considering a number of directions for future research. Here are two.

Composing Schneider's security automata is straightforward [Sch00], but this is not the case for our edit automata. Since edit automata are sequence transformers, we can easily define the composition of two automata E_1 and E_2 to be the result of running E_1 on the input sequence and then running E_2 on the output of E_1. Such a definition, however, does not always give us the conjunction of the properties enforced by E_1 and E_2. For example, E_2 might insert a sequence of actions that violates E_1. When two automata operate on disjoint sets of actions, we can run one automaton after another without fear

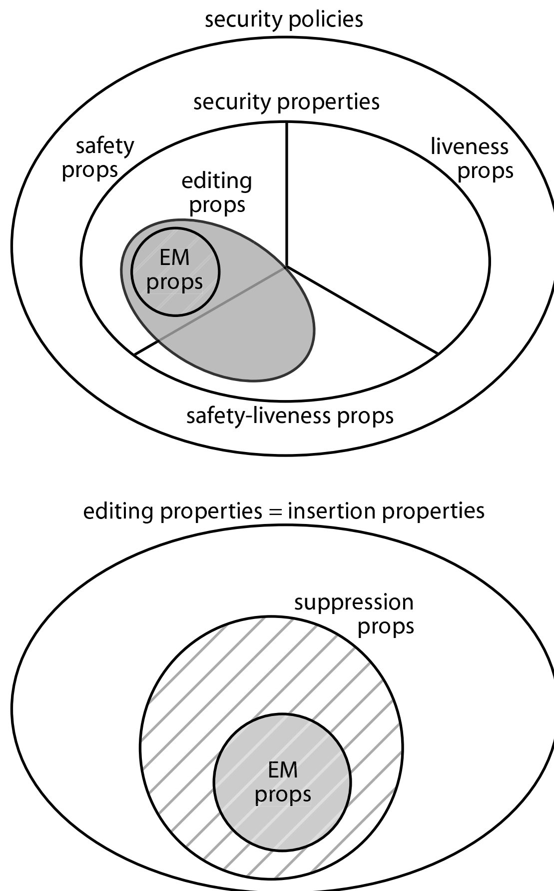

Figure 5: A taxonomy of *precisely* enforceable security policies.

that they will interfere with one other. However, this is not generally the case. We are considering static analysis of automaton definitions to determine when they can be safely composed.

Our definitions of precise and conservative enforcement provide interesting bounds on the strictness with which properties can be enforced. Although precise enforcement of a property is most desirable because benign executions are guaranteed not to be disrupted by edits, disallowing even provably benign modifications restricts many useful transformations (for example, the enforcement of the market policy from Section 3.4). Conservative enforcement, on the other hand, allows the most freedom in how a property is enforced because *every* property can be conservatively enforced by an automaton that simply halts on all inputs (by our assumption that $\hat{P}(\cdot)$). We are working on defining what

it means to *effectively* enforce a property. This definition may place requirements on exactly what portions of input sequences must be examined, but will be less restrictive than precise enforcement and less general than conservative enforcement. Under such a definition, we hope to provide formal proof for our intuition that suppression automata effectively enforce some properties not effectively enforceable with insertion automata and vice versa, and that edit automata effectively enforce more properties than either insertion or suppression automata alone.

6 Conclusions

In this paper we have defined two new classes of security policies that can be enforced by monitoring programs at runtime. These new classes were discovered by considering the effect of standard editing operations on a stream of program actions. Figure 5 summarizes the relationship between the taxonomy of security policies discovered by Alpern and Schneider [AS87, Sch00] and our new editing properties.

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