Automatic detection and demonstrator generation for information flow leaks in object-oriented programs

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ABSTRACT

We present a method to generate automatically exploits for information flow leaks in object-oriented programs. The goal, similar to white-box test generation, is to automatically produce executable, reusable test cases that challenge a given information flow policy with a very high degree of guaranteed coverage. Our approach combines self-composition and symbolic execution to create an insecurity formula for a given program and information flow policy. Satisfiability of this formula signifies the presence of information leaks and permits to use model generation for creating exploits. We support different kinds of information flow policies like noninterference, delimited information release, and information erasure. A prototypic tool implementation for Java programs of our approach is available. It generates exploits in the form of self-contained, executable JUnit tests. We evaluate our method and tool based on a set of micro-benchmarks and a case-study on e-voting.

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1. Introduction

Ensuring that programs do not leak unintended information is essential to establish confidence in the ability of an IT system not to put the security and privacy of its users at risk. In the past decade, much theoretical and practical work on information flow analysis of programs was produced (for example Askarov et al., 2015; Banerjee and Naumann, 2005; Barthe et al., 2004, 2011; Dam et al., 2010; Darvas et al., 2005; Devriese and Piessens, 2010; Hunt and Sands, 2006; Myers, 1999; Sabelfeld and Myers, 2004; Scheben and Schmitt, 2012; Volpano et al., 1996). Its focus is to ensure that an outside agent with well-defined properties cannot infer secret inputs by observing (and initiating) several runs of a program.

Much research on information flow analysis aims at statically proving that a program does not leak any unintended information. One can distinguish type-based (e.g. Banerjee and Naumann, 2005; Hunt and Sands, 2006; Myers, 1999; Sabelfeld and Myers, 2004; Volpano et al., 1996) and logic-based (e.g. Barthe et al., 2004; Darvas et al., 2005; Scheben and Schmitt, 2012) approaches. The former tends to be automatic, but imprecise, while the latter tends to be precise, but requires expert interaction.

Precision and automation at the same time can be achieved dynamically, for a concrete program run. Security monitors (e.g. Askarov et al., 2015; Dam et al., 2010) raise a warning as soon as a program violates a given security policy and try to contain the leak. Secure multi-execution (Devriese and Piessens, 2010) and several follow-up papers) determines whether a given concrete run might violate a policy.

In this paper we connect the static and dynamic view: by static analysis we produce test cases, which we call exploits, that are guaranteed to violate a given security policy. Such...
exploits can then be used independently for system testing, regression testing, documentation, etc. We take a similar view here as in functional verification: a security policy can be seen as a requirements specification. Its violation is a software fault witnessed by a test case, i.e. exploit.

In functional verification it is well-known that static verification is not a replacement for testing, but both techniques complement each other (Beckert and Hähnle, 2014): static verification gives strong guarantees for the modeled part of a system, while testing is incomplete, but validates actual executables and can also find problems in the underlying runtime environment or hardware platform. In our work we intend to achieve similar goals as in white-box software test generation frameworks (Albert et al., 2010; de Halleux and Tillmann, 2008; Engel and Hähnle, 2007).

- Under the assumption that a sound and complete specification is provided, no further user interaction is required and the approach is automatic.
- Completeness is achieved only in specific cases (e.g., no loops present or else strong loop invariants are supplied), however, strong and precise coverage guarantees can be given.
- Like test cases, exploits can be used to validate a program in its actual runtime environment.
- Like test cases, exploits become part of a library that is regularly executed to protect against regression. They are useful even after changes were made.
- Like test cases, exploits can serve as a documentation and illustration of intended system behavior.

Unlike monitoring and multi-execution, no special runtime infrastructure is required, because we generate exploits in the form of self-contained JUnit tests. These run the program under test multiple times with the generated input in order to produce a security violation. In addition, the generated exploits come with statically determined coverage guarantees.

A huge problem in software testing is the creation of oracles (Barr et al., 2015) that tell whether a test succeeded. Among the problems are i) missing or insufficient specifications, ii) complexity of general functional specifications which might contain quantifications over all objects and similar and iii) possible unintended side-effects of specification code (runtime exceptions and similar). Even in automatic test generation, these must often be supplied manually. But information flow policies (see Sabelfeld and Sands (2009) for an overview) can usually be expressed in a uniform manner for any given program. Hence, it is possible to generate test oracles automatically from them. In fact, our approach goes one step further in being oracle-sensitive: only such exploits that violate a given policy are generated at all.

Like other white-box software test generation frameworks (Albert et al., 2010; de Halleux and Tillmann, 2008; Engel and Hähnle, 2007; King, 1976), our method is based on symbolic execution of the target program. In our case, it is embedded (Glädisch, 2008) into the program logic of the deductive verification framework KeY (Beckert et al., 2007), which is also the basis for the acronym of our tool KEG (for KeY Exploit Generator). The relational nature of information flow (two runs of a program must be compared) is captured by the technique of self-composition (first introduced in Darvas et al. (2003); the name self-composition was coined in Barthe et al. (2004)). The result is an insecurity formula for a given information flow policy that is satisfiable if and only if the policy is violated. Model generation with the help of SMT solvers (de Moura and Björner, 2008) yields the input data for the exploit. In addition to standard information flow policies like noninterference, we also support relativized properties that tend to be used in practice, including delimited information release and information erasure (Sabelfeld and Sands, 2009).

The paper is structured as follows. Section 2 introduces basic notions and techniques. Section 3 explains the logic formalization of security policies. In Section 4 we discuss the analysis of loops and method invocations, needed to deal with realistic programs. Section 5 presents our tool KEG and gives a proof-of-concept with micro benchmarks. Our method and its implementation KEG are evaluated with the help of a larger case study (an e-voting program) in Section 6. We compare our work with others in Section 7 and conclude with Section 8.

This paper is an extended version of Do et al. (2015). The main new aspects are a generalization of the encoding of noninterference to include information erasure (Section 3.2) and the e-voting case study in Section 6 which is completely new. Minor additions concern micro benchmarks (Section 5.3) and a detailed description of the structure of the generated JUnit tests (Section 5.2).

2. Background

2.1. Information flow policies

Before we can analyze that a program does not leak confidential information, we need to define the security requirements. This has two aspects: the security level of each program location (i.e. program variables and fields) as well as an information flow policy defining whether and what kind of information may flow between program locations with a different security level. We recall the definitions of two well-known information flow policies supported by our approach.

Noninterference. Noninterference (Cohen, 1978; Volpano et al., 1996) is the strongest possible information flow policy. It typically involves two security levels (high/confidential vs. low/public) and completely prohibits any information flow from program locations containing confidential information) to publicly observable program locations. The opposite direction is allowed. In our work we consider only deterministic programs. In this case, noninterference can be formalized by comparing two program runs:

Definition 1. (Noninterference – Informal). A program has secure information flow with respect to noninterference, if any two executions of the program starting in initial states with identical values of the low variables, also end in final states which coincide on the values of the low variables.

In other words the final value of low variables is solely determined by the initial value of low variables and does not depend on the initial values of high variables.
We define some basic notions required to formalize information flow policies. Let \( p \) denote a program and \( \sigma \) the set of all program variables\(^1\) of \( p \).

**Definition 2. (Program State).** A program state \( \sigma \) maps each program variable \( v \in \text{Var} \) of type \( T \) (write \( v : T \)) to a value of its concrete domain \( D^T \), i.e.:

\[
\sigma : \text{Var} \rightarrow D
\]

with \( \sigma(v : T) \in D^T \) and \( D \) being the union of all concrete domains. The set of all states for a given program \( p \) is denoted as \( \text{States}_p \).

We define coincidence of program states relative to a set of program variables:

**Definition 3. (State Coincidence).** Given a set of program variables \( V \) and two states \( \sigma^1, \sigma^2 \in \text{States}_p \), we write \( \sigma^1 = \sigma^2 \) if and only if \( \sigma^1(v) = \sigma^2(v) \) for all \( v \in V \).

A concrete execution trace \( \tau \) of a program \( p \) is a possibly infinite sequence of program states \( \tau = \sigma_0, \sigma_1, \sigma_2, \ldots \) produced by starting \( p \) in state \( \sigma_0 \). In this paper, we concern ourselves only with terminating\(^2\) programs, consequently, all of our execution traces are finite. Then the big-step semantics is defined as follows: let \( X \) be a concrete execution of a program \( p \) defined by a trace \( \tau \). We represent \( X \) by a pair \( (\tau, \sigma) \), where \( \sigma \in \text{States}_p \) is the first state of \( \tau \) and \( \sigma_\text{out} \in \text{States} \) is the last. The set of all possible concrete executions of \( p \) is denoted as \( \text{Exc}_p \). We can now formally define noninterference for two security levels low and high:

**Definition 4. (Noninterference).** Given a program \( p \) over variables \( \text{Var} \) and a noninterference policy \( \text{NI} = H \rightarrow L \) where \( L \cup H = \text{Var} \) such that \( L \) contains the low variables and \( H \) the high variables, program \( p \) has secure information flow with respect to \( \text{NI} \) if and only if for all concrete executions \( X, Y \in \text{Exc}_p \) it holds that if \( \sigma^1 \downarrow = \sigma^2 \downarrow \) then \( \sigma^1_\text{out} \downarrow = \sigma^2_\text{out} \downarrow \).

**Example 1.** The program:

\[
\text{if} \ (h > 0) \ {\{1 - 2; \}}
\]

with high variable \( h \) and low variable \( 1 \) is insecure as it does not satisfy the noninterference property for the policy \( \text{NI} = [h] \rightarrow [1] \). Given two initial states \( \sigma^1 \) with \( \sigma^1(h) = 5 \), \( \sigma^1(1) = 0 \) and \( \sigma^2 \) with \( \sigma^2(h) = -5 \), \( \sigma^2(1) = 0 \), respectively. They satisfy \( \sigma^1 =_1 \sigma^2 \), but in the final states we have \( \sigma^1_\text{out} (1) = 2 \neq \sigma^2_\text{out} (1) = 0 \).

Declassification. In practice noninterference is too restrictive. For instance, a program that authenticates users with their login password leaks the information whether an entered password is correct, or take a database that may be queried for aggregated values like the average salary, but not for the income of an individual person.

Declassification is a class of information flow policies that allows one to express that some precisely specified confidential information may be leaked. The paper Sabelfeld and Sands (2009) provides an extensive survey of decalification approaches. Here we consider delimited information release as introduced in Sabelfeld and Myers (2004). Delimited information release is a decalification policy which allows one to specify what kind of information may be released. To this end, so called escape hatch expressions are specified in addition to the security level of the program locations. For instance, the escape hatch:

\[
\sum_{e \in \text{Excp}} \text{salary}(e) \left[ \frac{\text{Person}}{\text{E}} \right]
\]

can be used to declassify the average of the income of all persons in a database. The formal definition of delimited information release extends Definition 4. Both definitions coincide for trivial escape hatch expressions such as \( e = \text{true} \).

**Definition 5. (Delimited Information Release).** Given a program \( p \) over variables \( \text{Var} \) and a delimited information release policy \( \text{Dec} = (L, H, E) \) with \( L, H \) as before and \( E \) denoting a set of escape hatch expressions, program \( p \) has secure information flow with respect to \( \text{Dec} \) if and only if for all concrete executions \( X, Y \in \text{Exc}_p \) it holds that if \( \sigma^1 =_1 \sigma^2 \) and for all \( e \in E \): \( \text{e}_p^\in = \text{e}_p^\out \), then \( \sigma^1_\text{out} = \sigma^2_\text{out} \). The expression \( \text{e}_p^\in \) denotes the semantic evaluation of \( e \) in state \( \sigma \).

**Example 2.** Consider again the program from Example 1:

\[
\text{if} \ (h > 0) \ {\{1 - 2; \}}
\]

Given the delimited information release policy \( \text{Dec} = ([1], [h], [h > 0]) \) where the escape hatch allows the sign of \( h \) to be leaked, the counter example for noninterference from Example 1 is no longer a counter example as:

\[
[h > 0]_n^\out = \text{true} \neq \text{false} = [h > 0]_o^\out.
\]

In fact the program is secure for the given policy, as the decision whether \( 1 \) may be altered is only based on the guard. The same policy for the program:

\[
\text{if} \ (h > 0) \ {\{1 - h; \}}
\]

is insecure as can be demonstrated by the following counter example: Given initial states \( \sigma^1 \) with \( \sigma^1(h) = 3 \), \( \sigma^1(1) = 0 \) and \( \sigma^2 \) with \( \sigma^2(h) = 5 \), \( \sigma^2(1) = 0 \) we observe (i) both states coincide on the value of the low variable \( 1 \); and (ii) they evaluate the escape hatch expression to the same value \( [h > 0]_n = [h > 0]_o = \text{true} \), but their final states differ on the value of \( 1 \):

\[
\sigma^1_\text{out} (1) = 3 \neq \sigma^2_\text{out} (1).
\]

In Section 3.2 we will further generalize the delimited information release policy to include information erasure.

### 2.2. Logic-based information flow analysis

**Symbolic Execution.** Symbolic Execution (King, 1976) is a versatile technique used for various static program analyses.

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\(^1\) To keep the presentation manageable, in the formal definitions we mention only variables, however, our implementation works also for fields, including reference types.

\(^2\) This is purely in the interest of scoping. Our approach is not principally limited to terminating programs.
Symbolic execution of a program means to run it with symbolic input values instead of concrete ones. Such a run results in a tree of symbolic execution traces, which cover all possible concrete executions.

Each node in a symbolic execution tree is annotated by its symbolic state given as a tuple of symbolic expressions, each tuple element corresponding to a program variable. In the example shown in Fig. 1b, the root node is a branching node whose outgoing edges are annotated by their branch conditions. Here the symbolic execution splits into two branches: the left one for the case where the symbolic value \( x_0 \) is non-negative and the right one for a negative \( x_0 \). Both branches might be taken as we do not have any further information about the value of \( x_0 \). The path condition of a path is the conjunction of all its branch conditions and characterizes the symbolic execution path uniquely. As long as the program does not contain loops or method invocations, a path condition is a quantifier-free formula in first-order logic.

From the tree in Fig. 1b we can extract that in case of a non-negative input value for \( x_0 \), the program terminates in a final state in which the final value of \( x_0 \) remains unchanged (i.e., \( x_0 \)) while the final value of \( y \) is \( 2(y_0 - 1) \). We make some notational conventions: Given path \( i \) we refer to its path condition by \( pc_i \) and to the final value of a program variable \( v \) on that path with \( f_v^i \). If we want to make explicit that the final value of a program variable \( v \) depends on the symbolic input value of a program variable we pass it as an argument to \( f_v \). For instance, if the left branch is numbered with 0, then the final value of \( y \) on that branch is \( f_y^0(x_0, y_0) = 2(y_0 - 1) \).

In case of unbounded loops or unbounded recursive method calls a symbolic execution tree is no longer finite. We overcome this obstacle and achieve a finite representation by making use of specifications as proposed in Hentschel et al. (2014). That approach uses loop invariants and method contracts to describe the effect of loops and method calls. The basic idea is that loop invariants and method contracts contribute to path conditions and to the modification of the symbolic state. This is realized in the symbolic execution engine of the verification system KeY (Beckert et al., 2007), which we use as implementation basis of KEG.

Self-composition. Our exploit generation approach is derived from a logic-based formalization of noninterference using self-composition as introduced in Darvas et al. (2003, 2005), based on a direct semantic encoding of noninterference in a program logic. The Hoare triple \( \{ \text{Pre} \} p \{ \text{Post} \} \) characterizes that, whenever the program \( p \), started in an initial state satisfying Pre, terminates, then Post must hold in the final state reached. Non-interference as given in Definition 4 requires the comparison of two program runs. In Darvas et al. (2005) this is achieved by copying the program \( p \) and replacing all its variables with fresh copies, such that the original and the copied version do not share any memory. Concretely, let without loss of generality \( l \in L, h \in H \) be the only variables of \( p = p(l, h) \). Further, let \( p(l', h') \) represent the copied program constructed from \( p \) by renaming variable \( l \) to \( l' \) and \( h \) to \( h' \). Then:

\[
(1 \equiv l') p(l, h); \quad p(l', h') \quad (1 \equiv l')
\]

is a direct formalization of noninterference. A major drawback of this formalization is that it requires program \( p \) to be analyzed twice. Several refinements have been presented to avoid the repeated execution (Barthe et al., 2011; Terauchi and Aiken, 2005). Here we use a different approach to that problem based on symbolic execution. The fundamental idea is to execute the program symbolically only once and then to use the path conditions and symbolic states to construct a single first-order formula with that has same meaning as the Hoare triple. To express the noninterference property, it is then sufficient to copy path conditions and symbolic values, replacing the symbolic input values with fresh copies. The technical details are given in the following section.

3. Exploit generation for insecure programs

We describe the formalization in logic for various information flow policies. Given a complete symbolic execution tree for a program and an information flow policy, we construct formulas that are unsatisfiable when the program is secure and satisfiable if the policy can be violated by some inputs. This approach permits to use an SMT solver or model finder to search for satisfying models from which one can then read off concrete input states for two program runs that demonstrate a violation of the given policy, see also Fig. 2.

3.1. Logic characterization of insecurity

First we show how to construct a formula that characterizes noninterference (Definition 4) from a complete symbolic
execution tree for program $p$ with paths $i \in \{0, \ldots, n-1\}$. Let $pc(l, H)$ be the path condition that uniquely determines path $i$. Let $NI = H \Rightarrow L$ be the noninterference policy with low variables $L$ and high variables $H$.

To represent two independent program runs, we create a copy of all program variables $Var' = \{v' \mid v \in Var\}$ with the notational convention that $v'$ always refers to the copy of $v$. Now we obtain the sets $L'$ and $H'$ as copies of $L$ and $H$, i.e., $L' = \{l' \mid l \in L\}$ (analogously $H'$). To refer to the initial (symbolic) value of a program variable $v$, instead of introducing a new constant symbol $v_0$, we simply use the program variable $v$ itself. Intuitively, the first run is performed using $Var$, while the second one uses the copy $Var'$. Both runs are independent as they do not share any common memory.

Then the NI-insecurity formula:

$$\forall i, j \in \text{nic} . \left( \langle l' \mid l \in L' \rangle \land pc(l, H) \land pc(l', H') \land \forall i, j . \left[ f(L, H) \neq f'(L', H') \right] \right)$$

is satisfied if and only if there is a model (i.e., concrete state) $\sigma$ assigning values to the program variables $Var$, $Var'$ such that

- the input values of the low variables $L$ coincide with the values of their copies in $L'$,
- there are two paths $i, j$ ($l = j$ possible) with consistent path conditions (i.e., both paths can actually be taken), but
- for which the final value of at least one low variable differs.

In other words, the model $\sigma$ assigns concrete values to $Var$ and $Var'$ such that $p$ produces different low level outputs for two runs from initial states with identical low input.

**Example 3.** The insecurity formula (1) for the example program from Fig. 1 and the NI policy $\{y \Rightarrow \{x\}\}$ is:

$$y_0 \equiv y_0' \land x_0 \geq 0 \land x_0' \geq 0 \land 2(y_0 - 1) \neq 2(y_0' - 1)$$
$$\lor y_0 \equiv y_0' \land x_0 \geq 0 \land x_0' < 0 \land 2(y_0 - 1) \neq 2(y_0' + 1)$$
$$\lor y_0 \equiv y_0' \land x_0 < 0 \land x_0' < 0 \land 2(y_0 + 1) \neq 2(y_0' + 1)$$

It is easy to see that the first and third disjuncts are unsatisfiable, but the second disjunct is satisfiable, e.g., for the model $x_0 \rightarrow 0$, $x_0' \rightarrow -1$, $y_0 \rightarrow 1$, $y_0' \rightarrow 1$. □

The NI-insecurity formula (1) can be rewritten into the equivalent formula:

$$\forall i, j \in \text{nic} . \left( \langle l' \mid l \in L' \rangle \land pc(l, H) \land pc(l', H') \land f(L, H) \neq f'(L', H') \right)$$

This formulation will be easier to incorporate declassification. The intuition behind the formula $\text{Leak}^NI(H, L, i, j)$ is that it allows us to ascribe leaks to a specific target, i.e., it is satisfiable, if some information is leaked from the program variables in $H$ to variable $l$.

The copy of the low level variables is actually not needed (as we require their equality for all models), so formulas (1) and (2) can be made more succinct by replacing $L'$ with $L$ and omitting the first conjunct, which states $L \equiv L'$. In the future we will tacitly perform this simplification. Then the first disjunct in Example 3 becomes $x_0 \geq 0 \land x_0' < 0 \land 2(y_0 - 1) \neq 2(y_0' - 1)$.

### 3.2. Generalized noninterference policy

Sometimes it is not sufficient to simply ensure that no information is leaked, but one wants also to guarantee that secret data are not kept longer than needed, because of legal reasons or to make data dumps (initiated by an attacker) less useful. Information erasure is for a desired property for cryptographic devices (secret keys must be erased after usage), online transactions (credit card information must be erased after the transaction is completed), e-voting (all data connecting voter and ballot must be erased after the result has been published), etc. Information erasure policies have been presented in Del Tedesco et al. (2011) and Hunt and Sands (2008).

**Listing 1.1: Ticket vending machine**

```c
void buy() {
    charge(ticketCost, ccNumber);
    log();
    ccNumber = 0;
}
```

**Example 4.** Consider a simple ticket vending machine model (adapted from Del Tedesco et al. (2011)) as shown in Listing 1.1. Assume that before executing the program the buyer’s credit card number was read from a terminal and stored in variable $ccNumber$. To complete the purchase, method `buy()` is called and the card account debited. After logging the purchase, the credit card number is erased from memory by setting $ccNumber$ to 0. This ensures that even a powerful attacker who can dump the memory of the vending machine to read the location of variable $ccNumber$ cannot learn anything about the credit card number of the buyer.

In more technical terms, we want to ensure that after execution of `buy()` the value of $ccNumber$ is permitted to flow to any low location. This policy is not expressible within the standard noninterference framework. A naive and ad hoc extension would be to classify $ccNumber$ as high and to add a constraint $f_{ccNumber}(L, H) \neq 0$ to the insecurity formula. But this does not work, for instance, if we want to erase the secret with a random number. □
To provide support for some of information erasure policies (Del Tesesco et al., 2011; Hunt and Sands, 2008) we generalize our notion of interference (cf. Definition 4).

Definition 6. (Generalized Noninterference). Given a program \( p \) over variables \( \text{Var} \), a generalized noninterference policy (GNI) is an ordered pair, written as \( H \vdash \text{GNI} L \), where \( L, H \subseteq \text{Var} \) are sets of low and high variables. Program \( p \) has secure information flow with respect to GNI if and only if for all concrete executions \( X, Y \in \text{Exce}_p \) it holds that if \( \sigma^X =_{\text{Var} \cup H} \sigma^Y \) then \( \sigma^X_{\text{out}} = \sigma^Y_{\text{out}} \).

The definition omits the requirement that \( L \) and \( H \) form a partitioning of \( \text{Var} \), i.e. a variable \( v \) is allowed to be a member of both variable sets. In addition, it potentially strengthens the condition on the output values of the low variables in the final states. The output values of the low variables in the final states must now be identical for any two initial states that coincide on the variable set of \( \text{Var} \cup H \). The set \( \text{Var} \cup H \) might not contain all variables in \( L \) (e.g. variable \( v \) from before would not be in \( H \)) and hence allows for more pairs of initial states to be considered. Generalized noninterference reduces to standard noninterference when \( L \cup H = \text{Var} \).

Example 5. (Example 4 Continued). To ensure that the credit card number is erased we specify the GNI policy.

\[ H = \{ \text{ccNumber} \} \vdash \text{GNI} \{ \text{ccNumber}, \text{ticketCost} \} = L \]

To keep the analysis simple, assume that methods charge (int, int) and log() do nothing. First we assume that line 4 that performs the forgot was ignored. The following test case demonstrates a violation of the policy: Given initial states \( \sigma_i, \sigma_j \):

\[
\begin{array}{c|c|c}
\text{ccNumber} & \text{ticketCost} \\
\hline
\sigma_1 & 1234 & 50 \\
\sigma_2 & 5678 & 50 \\
\end{array}
\]

Both initial states coincide on the variable set \( \text{Var} \setminus H = \{ \text{ticketCost} \} \). To adhere to the GNI policy, the final states \( \sigma^i_{\text{out}}, \sigma^j_{\text{out}} \) must coincide on all low variables, i.e. on the variables ticketCost and ccNumber. But as none of the values is changed by the runs, they still differ in the value of ccNumber, hence, the GNI policy is not valid. Now assume line 4 is present; then, in both executions the final value of ccNumber is 0 and the GNI property holds.

The logic formalization of the corresponding insecurity formula is almost identical to (2):

\[
\forall i, j \in \text{Loc} \cdot (\text{ccNumber} \setminus H) \cdot \left( \forall v \in \text{Var} \setminus H \cdot (v \not\in \text{Var}) \land p_C(\text{Var}) \land p_C(\text{Var}') \land f_i(\text{Var}) \not= f_j(\text{Var}') \right)
\]

3.3. Targeted conditional delimited release

We further extend the insecurity formula for generalized noninterference (3) to delimited information release (DIR) (Sabelfeld and Myers, 2004). In contrast to the standard version of DIR, our policy describes not only what information can be released by escape hatches, but also allows to express under which condition and to whom (target) the information might be leaked.

Definition 7. (Targeted Conditional Delimited Release). Given a program \( p \) over variables \( \text{Var} \) and a GNI \( H \vdash \text{GNI} L \). A Targeted Conditional Delimited Release (TCD) policy \( (D, \text{GNI}) \) is a set of specification triples where each \( (e, C, T) \in D \) consists of:

- an escape hatch expression (i.e. first-order term) \( e \) over \( \text{Var} \),
- a declassification condition formula \( C \) over \( \text{Var} \), and
- \( T \subseteq L \), a set of program variables to which the specified escape hatch is allowed to be leaked.

A program satisfies a given TCD policy \( (D, \text{GNI}) \) if it satisfies the GNI policy, except for the cases covered by a triple \( (e, C, T) \in D \). Here, program is free to release the information captured by the escape hatch expression \( e \) to a location in \( T \) provided that condition \( C \) is satisfied in the initial state of the execution.

Given a TCD policy \( (D, \text{GNI}) \) and a program \( p \). We give the insecurity formula for the case that \( D = \{(e, C, T)\} \) consists of a single TCD specification triple:

\[
\forall i, j \in \text{Loc} \cdot \text{Loc}_{\text{ccNumber}}(H, L, i, j) \land \left( (\exists t \in T \cdot C(\text{Var}) \land C(\text{Var}') \rightarrow e(\text{Var}) \not= e(\text{Var}')) \right)
\]

The formula coincides with the noninterference insecurity formula for locations \( l \notin T \) that are not among the allowed release targets. Otherwise, the new second conjunct adds:

\[
C(\text{Var}) \land C(\text{Var}') \rightarrow e(\text{Var}) = e(\text{Var}')
\]

as an additional restriction to the initial states for both runs: if both initial states satisfy the declassification condition \( C \) then they must also coincide on the value of the escape hatch expression. The justification is that if there are two runs such that their initial states coincide on the low level input and on the escape hatches and if the final value for the allowed target differs, then more information than just the escape hatch must have been released.

Example 6. (Example 2 Continued). Consider the program from Example 2:

\[
\text{if } (\text{h} > 0) \{ \text{i = 2} \}
\]

Let \((\{(\text{h} > 0, \text{true}, \{1\})\}, \{\text{i} \} \vdash \text{GNI} \{\{\text{i}\}\})\) be a declassification TCD policy. This is the insecurity formula resulting from symbolic execution starting in \((\text{i} = 0, \text{h} = 0)\):

\[
(\text{i} = 0 \land \text{i} > \text{h} > 0 \land \text{h} > 0 \land \text{h} > 0 \land \text{h} > 0) \lor
(\text{i} = 0 \land \text{i} > \text{h} > 0 \land \text{h} > 0 \land \text{h} > 0) \lor
(\text{i} = 0 \land \text{i} > \text{h} > 0 \land \text{h} > 0 \land \text{h} > 0) \lor
(\text{i} = 0 \land \text{i} > \text{h} > 0 \land \text{h} > 0 \land \text{h} > 0) \lor
(\text{i} = 0 \land \text{i} > \text{h} > 0 \land \text{h} > 0 \land \text{h} > 0)
\]

The first and third disjuncts are trivially invalid. The second disjunct is invalid, because the second conjunct implies \(\text{h}_0 > 0\)
if and only if \( h' > 0 \) which contradicts the path condition in the first conjunct. Consequently, the insecurity formula is unsatisfiable, i.e. the program is secure for the specified policy. Consider a slightly altered program:

\[
\text{if } (h > 0) \{ l + 2; \}
\]

We analyze it with the same policy and initial state as above. Now in the resulting insecurity formula:

\[
(l_0 \equiv l' \wedge h_0 \geq 0 \wedge h'_0 \geq 0 \land 2 \neq 2) \wedge ((\text{true} \land \text{true}) \rightarrow (h_0 > 0) \wedge (h'_0 > 0)) \vee
(l_0 \equiv l' \land h_0 < 0 \land h'_0 \geq 0 \land l_0 \neq 2) \wedge ((\text{true} \land \text{true}) \rightarrow (h_0 > 0) \wedge (h'_0 > 0)) \vee
(l_0 \equiv l' \land h_0 < 0 \land h'_0 < 0 \land l_0 \neq 2) \wedge ((\text{true} \land \text{true}) \rightarrow (h_0 > 0) \wedge (h'_0 > 0))
\]

the second disjunct is satisfiable, for instance, when the initial value of \( h \) is \(-1\) and the initial value of \( h' \) is \( 0 \). Consequently, the program does not adhere to the specified policy.

4. Exploit generation using program specifications

Code with unbounded loops or recursive method calls gives rise to infinite symbolic execution trees. Another difficulty is posed by calls to library methods for which no source code is available. And in general symbolic execution trees tend to become infeasibly large when the implementations of called methods are simply inlined. To overcome these problems we annotate programs with specifications in the form of loop invariants and method contracts. This allows to approximate a specification during symbolic execution and we can adapt that approach accordingly.

4.1. Loop specification

To compute the path conditions and the final values of symbolic execution paths we need to be able to execute unbounded loops without unwinding them infinitely often. In program verification this is achieved by providing a loop specification. A loop specification \( LS = (I, \text{mod}) \) consists of a loop invariant \( I \) and a set of program variables \( \text{mod} \) that contains at least those program variables the loop can possibly modify.

We need to integrate loop specifications into the NI-insecurity formula (2). Let \( b \) be the guard of a loop and \( LS = (I, \text{mod}) \) its specification. The basic idea in Hentschel et al. (2014) is that the loop specification describes the state after exiting the loop. This means, we can treat the loop as a black-box and continue execution after the loop in a state for which the variables \( \text{mod} \) that might have been modified by the loop are set to an unknown value. Unknown values are represented by fresh symbolic values \( V_{\text{mod}} \). The only knowledge about these values is provided by the loop invariant and by the fact the loop guard \( b \) must be false after exiting the loop.

Our insecurity formulas express a constraint over the initial state. For instance, the final value \( f' \) of variable \( l \) is given in terms of the initial symbolic values of the program variables. The same holds for the path conditions. We make this implicit weakest precondition computation here explicit for the loop guard and the invariant, i.e. \( f_{\text{wp}} \) is the weakest precondition of \( l \) computed in the state directly after the loop (similar for the loop guard).

For the sake of simplicity, we only show how to adapt \( \text{Leak}^{\text{NI}}(H, l, i, j) \) for the case that both paths \( i, j \) contain the same loop:

\[
\text{Leak}^{\text{NI}}(H, l, i, j) = \left( \land_{\text{inv}} (V) \land \text{pc}(V) \land \text{pc}(V_{i}) \right) \land \text{Pre}(V) \land \text{Post}(V_{i}) \land f(V_{i}) = f(V)
\]

Here \( V_{i} = \text{Var} \cup V_{\text{mod}} \) and \( h_{\text{wp}}(V_{i}) \) are the symbolic values of the guard after the loop expressed in terms of the initial values of \( V_{i} \). If one or both of paths \( i, j \) do not contain the loop or a different loop, then the conjuncts corresponding to the invariants and loop guards are omitted or added accordingly.

Example 7. We illustrate formula (6). Consider the loop below with low variable \( l \) and high variable \( h \). We want to establish whether this code is secure with respect to the policy \( \{ h \} \bowtie \{ \} \).

\[
\begin{align*}
\text{while } (l > 0) \{ l = l - 1; \} \\
1 = h \ast h; \\
\text{while } (1 > 0) \{ l + 1; \} \\
1 = l + h; \\
\end{align*}
\]

Let \( l_{\text{mod}}, l'_{\text{mod}} \) be the fresh values representing the value of \( l \) directly after the loop. Computing the weakest precondition of the invariant gives us \( l'_{\text{mod}} \geq 0 \) and for the guard \( l_{\text{mod}} > 0 \) for the first run (analogous for the second run). The resulting formula is (note that there is only one path and no path condition):

\[
(l_0 \equiv l' \land l_{\text{mod}} \geq 0 \land \neg(l_{\text{mod}} < 0) \land l_{\text{mod}} \\
\geq 0 \land \neg(l_{\text{mod}} > 0) \land l_{\text{mod}} + h_0 \neq l'_{\text{mod}} + h_0)
\]

The formula is satisfiable, for example, with \( l_0 = l' = 10, l_{\text{mod}} = l'_{\text{mod}} = 0, h_0 = 1 \) and \( h'_0 = 2 \). Indeed, the program is insecure. Removing the final statement would make it secure. In this case the final conjunct in the insecurity formula would change to \( l_{\text{mod}} \neq l'_{\text{mod}} \) which renders it unsatisfiable. \( \square \)

4.2. Method contracts

Let \( m \) be a method name. A contract \( C_{m} \) for \( m \) is a triple \((\text{Pre}_{m}, \text{Post}_{m}, \text{Mod}_{m})\) with precondition \( \text{Pre}_{m} \), postcondition \( \text{Post}_{m} \) and
modifies (or assignable) clause Mod. The latter is the set of all program variables whose value \( m \) can possibly change (similar as \( mod \) in loop specifications).

A method satisfies its contract, if it ensures that when invoked in a state for which the precondition is satisfied, then in the final state the postcondition holds and at most the value of program variables in the assignable clause has been modified.

Analyzing Noninterference Relative to a Precondition. Given a method \( m \) with contract \( C_m \). We want to analyze whether \( m \) respects a noninterference policy \( H \nleftrightarrow L \) under the condition that \( m \) is only invoked in states satisfying its precondition \( Pre_m \). Adapting the noninterference formula (2) is straightforward and merely requires to add a restriction to the initial states that they must satisfy the method’s precondition:

\[
\text{Leak}^{NI}(H, L, I, I, J) \land Pre_m(Var) \land Pre_m(Var')
\]

Analyzing Programs with Method Contracts for Noninterference. This problem is solved in Hentschel et al. (2014) by using method contracts in a similar way loop specifications have been used. Instead of a loop invariant, the pre- and postconditions become part of the path conditions. The modifies clause again introduces fresh values to represent the symbolic value of program variables that might have been changed as side effect of method invocation.

Let \( n \) be the method analyzed for secure information flow and assume it invokes method \( m \). Let the contract of \( n \) be \((Pre_n, Post_n, Mod_n)\). In the case that each of the paths \( i, j \) contains exactly one method call for \( n \) we obtain:

\[
\text{Leak}^{NI}(H, L, I, I, J) = \bigwedge_{l \in \mathbb{L}_i} v = v' \land pc_i(V_s) \land pc_i(V_s') \\
\land Pre_{n'}(V_s) \land Post_{n'}(V_s') \land Pre_{n'}(V_s') \\
\land Post_{n'}(V_s') \land f'(V_s) \neq f'(V_s') \tag{7}
\]

where \( V_s = \text{Var} \cup \text{Var}_{\text{Mod}} \) (analogous for the copies) and \( Pre_{n'} \), \( Post_{n'} \) are the weakest preconditions of \( Pre_n \), \( Post_n \) computed directly before and after method invocation, respectively. If method \( n \) returns a value, a fresh variable representing the return value of \( n \) is added to \( V_s \). The return value can be referenced in \( Post_n \). The general case (more method calls, or more than one method call) is handled similarly as in the loop case.

Example 8. We illustrate formula (7) with method \( \text{run}() \) shown in Listing 1.2. We want to establish whether \( \text{run}() \) is secure with respect to the policy \( \{ h \} \prec \{ \} \). We expect that it is insecure: the returned value of method \( \text{calc}(\text{int}) \) depends on its parameter. In line 3, \( h \) is passed as argument, hence, the returned value depends on high input, but it is assigned to low variable \( l \).

We construct the insecurity formula: method \( \text{run}() \) invokes the recursive method \( \text{calc}(\text{int}) \). To analyze the information flow resulting from this invocation, we have used a method contract for \( \text{calc}(\text{int}) \), because the recursion does not have a fixed bound.\(^3\) Let \( \text{calc}(\text{int}) \)'s contract be given as follows:

\[
\text{Pre}_\text{calc}: \; \text{true}
\]

\[
\text{Post}_\text{calc}: \; (x \leq 0 \rightarrow \text{result} = 0) \land (x > 0 \rightarrow 2 \times \text{result} = x \times (x + 1))
\]

\[
\text{Mod}_\text{calc}: \; \emptyset
\]

where \( \text{result} \) refers to the return value and \( \text{Mod}_\text{calc} \) is empty as \( \text{calc}(\text{int}) \) does not change the state.

Let \( r \) be a program variable representing the return value of \( \text{calc}(\text{int}) \). To apply the contract for the invocation at line 3, we need to instantiate the above contract as follows:

\[
\text{Pre}^{\text{calc}}_\text{run}: \; (h + 2 \leq 0 \rightarrow r = 0) \land (h + 2 > 0 \rightarrow 2 \times r = (h + 2) \times ((h + 2) + 1))
\]

\[
\text{Post}^{\text{calc}}_\text{run}: \; (h + 2 \leq 0 \rightarrow r = 0) \land (h + 2 > 0 \rightarrow 2 \times r = (h + 2) \times ((h + 2) + 1))
\]

\[
\text{The insecurity formula is then:}
\]

\[
\text{Leak}^{NI}(H, L, I, I, I) = \bigwedge_{l \in \mathbb{L}_i} v = v' \land pc_i(V_s) \land pc_i(V_s') \\
\land Pre_{n'}(V_s) \land Post_{n'}(V_s') \land Pre_{n'}(V_s') \\
\land Post_{n'}(V_s') \land f'(V_s) \neq f'(V_s')
\]

\[
\text{The formula is satisfiable, for example, with } h = 10, h_0 = 1, h_0' = 2, r_0 = 6 \text{ and } r_0' = 10 \text{ which means that method } \text{run}() \text{ is insecure.} \]

Listing 1.2: Recursive method call

```
1  public void run() {
2      h = h + 2;
3      l = calc(h);
4  }
5
6  private int calc (int x) {
7      if (x <= 0) return 0;
8      else return x + calc(x-1);
9  }
```

4.3. General observations and remarks

Using loop specifications or method contracts has one major drawback, namely, that not all models of a formula give rise to an actual information leak, or even worse, the insecurity formula of a secure program might become satisfiable. This case does not affect the soundness, but triggers false warnings. The reason is that the specifications might be too weak and allow behaviors that are not possible in the actual program. These false warnings can be filtered out by actually running the generated exploit. If the exploit fails to demonstrate the information leak, we know that our model was a spurious one. We can even start a feedback loop with a conflict clause which rules out the previously found model.

On the other hand, if loop or method specifications are not only too weak, but wrong in the sense that they exclude possible behavior, then leaks might go undetected. As we are concerned with bug detection and not verification, this is not too serious as we do not claim to find all bugs. Nevertheless,
5. Implementation and experiments

5.1. The KeY exploit generation tool and process

Our tool KeY Exploit Generation (KEG) implements the approach presented in the previous sections. KEG uses the symbolic execution engine of the verification system KeY (Beckert et al., 2007) for the analysis of Java programs. It supports method and loop specifications as explained in Section 4 to compute finite, yet complete (up to the specification) symbolic execution trees. The SMT solver Z3 (de Moura and Bjørner, 2008) is used to find models for the computed insecurity formulas. KEG fully supports Java reference types (objects and arrays). This makes it necessary to test the value of variables that have a reference type for equality. For this we use the approach proposed in Beckert et al. (2014). Comprehension expressions (max, min, 2) are also supported and can be used in specifications, including escape hatches, loop specifications, and method contracts.

Fig. 2 outlines KEG’s workflow. The input is a Java method m together with the information flow policy to which it has to adhere. Possibly, m is annotated with loop specifications. Contracts of methods called by m are provided if they are not to be inlined during symbolic execution. KEG expects the information flow policy to be present in the source code inside specifically marked Java comments. First, method m is symbolically executed (using KeY) to obtain a complete symbolic execution tree which can then be queried for the method’s path conditions and the final symbolic values of the program locations modified by m. Second, using the path conditions and symbolic final values, instances of the insecurity formulas described in Section 3 and 4 are generated. In a third step, these formulas are passed to a model finder (currently we use the SMT solver Z3 (de Moura and Bjørner, 2008)). If a model for the insecurity formula has been found, that model is used to determine the insecurity formulas is satisfiable by the following model provided by Z3:

The model describes two runs: The first run, labeled with 1, starts in an initial state with x, y and l initialized with 1, 1 and 0, respectively. We identify variable v in execution X by v.X. The second run, labeled with 2, has initial values −1, −1, and 0. The first run enters the then-branch of the conditional, the second one does not.

As the value of l is altered in the first run when executing the then-branch but not by the second run, where it remains 0, a leak is detected (the leaked information is the sign of field x). KEG generates exactly one exploit, which is output as a well-structured and human readable JUnit test. The exploit program is depicted in Listing 1.4. KEG outputs the exploit exactly as shown, i.e. pretty printed and structured with comments. We have only renamed a few fields to increase readability further.

The two initial states for the two runs of method magic are set up in lines 6–14 and 23–31. To ensure that the runs do not interfere, two instances of class Simple are created (lines 6 and 23). In general, more work needs to be invested to ensure independent runs (for instance, if static members are modified by a run). This can be achieved by running the program on two different Java Virtual Machines and by querying those for the required information. (This is currently not supported by KEG.)

Before invoking method magic, the initial values of all fields and method parameters are assigned. In Listing 1.4, the initial values of each field x, y of both runs are stored in the corresponding variables 1, 1, x_1, y_1 (lines 7–9) and 1, 2, x_2, y_2 (lines 24–26). These initial values are taken from the counter example produced by the SMT solver. To assign values to the

Listing 1.3: Example program to illustrate KEG

```java
public class Simple {
  public int m;
  private int x, y;
  // x ! | | x y ; /=
  // escapes (x*y) | to | if x=-1; @p/
  public void magic() { if (n=0) { l=x*y; } else { l=0; } }
}
```

---

incompleteness can be avoided by verifying the specifications using a program verification tool such as KeY (Beckert et al., 2007).

5.2. Exploit generation by example

We illustrate the work flow of KEG step by step with the program in Listing 1.3. Class Simple declares three integer typed fields x, y as well as a method called magic() which assigns a value to l depending on the sign of field x.

The information flow policy of the class is \( x, y \not\leftrightarrow \{1\} \) and specified in a comment starting with “/*!” in Line 4. Variables x, y are implicitly declared as high variables and l as a low variable. This strict noninterference policy is relaxed in line 6 for method magic() by providing a targeted conditional release specification consisting of an escape hatch \( x*y \), the target l and the condition \( x > -1 \).

Running KEG on the above example produces a symbolic execution tree consisting of two paths: one for each branch of the conditional statement. KEG generates for each unique pair of these paths the corresponding insecurity formulas and passes these on to the SMT solver. Only one of the three generated insecurity formulas is satisfiable by the following model provided by Z3:

<table>
<thead>
<tr>
<th>Insecurity formula (in SMT-LIB syntax)</th>
<th>Model</th>
</tr>
</thead>
<tbody>
<tr>
<td>( {\text{l} = \langle \langle x_1 \text{l} \rangle \text{not} \langle \langle x_2 \text{l} \rangle \text{x} \rangle \langle \langle x_2 \text{l} \rangle \rangle }} )</td>
<td>( x_1: 1 )</td>
</tr>
<tr>
<td>( \langle \langle x_2 \rangle \langle \langle x_2 \rangle \rangle } } } )</td>
<td>( x_2: -1 )</td>
</tr>
<tr>
<td>( \langle \langle y_1 \rangle \langle y_2 \rangle \rangle } } } } )</td>
<td>( y_1: 1 )</td>
</tr>
<tr>
<td>( \langle \langle y_1 \rangle \langle y_2 \rangle \rangle } } } } )</td>
<td>( y_2: -1 )</td>
</tr>
<tr>
<td>( \langle \langle l_1 \rangle \langle l_2 \rangle \rangle } } } } )</td>
<td>( l_1: 0 )</td>
</tr>
<tr>
<td>( \langle \langle l_2 \rangle \rangle } } } } )</td>
<td>( l_2: 0 )</td>
</tr>
</tbody>
</table>

The model describes two runs: The first run, labeled with 1, starts in an initial state with x, y and l initialized with 1, 1 and 0, respectively. We identify variable v in execution X by v.X. The second run, labeled with 2, has initial values −1, −1, and 0. The first run enters the then-branch of the conditional, the second one does not.

As the value of l is altered in the first run when executing the then-branch but not by the second run, where it remains 0, a leak is detected (the leaked information is the sign of field x). KEG generates exactly one exploit, which is output as a well-structured and human readable JUnit test. The exploit program is depicted in Listing 1.4. KEG outputs the exploit exactly as shown, i.e. pretty printed and structured with comments. We have only renamed a few fields to increase readability further.

The two initial states for the two runs of method magic are set up in lines 6–14 and 23–31. To ensure that the runs do not interfere, two instances of class Simple are created (lines 6 and 23). In general, more work needs to be invested to ensure independent runs (for instance, if static members are modified by a run). This can be achieved by running the program on two different Java Virtual Machines and by querying those for the required information. (This is currently not supported by KEG.)

Before invoking method magic, the initial values of all fields and method parameters are assigned. In Listing 1.4, the initial values of each field x, y of both runs are stored in the corresponding variables 1, 1, x_1, y_1 (lines 7–9) and 1, 2, x_2, y_2 (lines 24–26). These initial values are taken from the counter example produced by the SMT solver. To assign values to the
fields of an object, the auxiliary method `setFieldValue` makes use of Java’s reflection framework.

After each run (at line 17 and line 34), the concrete output value of `l` is extracted (lines 20 and 37). Line 45 asserts that the output values of `l` observed at the end of each run are equal. The JUnit test throws an assertion failure exception, if the output values of the two runs are different and thus an actual leak happened. If no exception is thrown then the counter example was spurious, for instance, due to a too weak loop specification. Tests that do not result in an assertion failure can then be omitted from our test suite.

5.3. Experiments

We performed experiments on micro benchmarks as a first part of the evaluation of our approach. Table 1 shows the aggregated results. All experiments were done on an Intel Core i7-4702HQ processor with JVM setting `-Xmx4096m`.

Concerning the runtime performance: A significant amount is spent for parsing the program, this can be reduced by parser optimizations, for example, by using a hand-coded version instead of a generated parser. Model finding time can be further optimized by performing simple techniques like symmetry reduction, learning and caching, all of which have not yet been implemented. Another factor is the programming language Java whose optimizations are performed at runtime and, hence, code that is run only few times will not be optimized at all.

A few observations concerning the benchmarks: For the examples `Mul` and `Comp`, we analyzed the effect of loop and method specifications in case of strong, weak and wrong specifications (`filename_Strong/Weak/Wrong_LI/MC`). As expected, with sufficiently strong specifications, all insecure paths could be precisely identified and only actual exploits were generated. Weak specifications over-approximate the behavior, leading to false positives, while wrong specifications can prevent to analyze all possible

---

Table 1 – Benchmark statistics.

<table>
<thead>
<tr>
<th>File name</th>
<th>Analyzed method</th>
<th>#L/MI</th>
<th>Policy (NI/D)</th>
<th>S/I</th>
<th>T_L (ms)</th>
<th>T_IX (ms)</th>
<th>T_MF (ms)</th>
<th>T_Tot (ms)</th>
<th>#GE/FW</th>
</tr>
</thead>
<tbody>
<tr>
<td>Mul</td>
<td>product</td>
<td>0/0</td>
<td>D</td>
<td>I</td>
<td>4187</td>
<td>847</td>
<td>1188</td>
<td>6266</td>
<td>1/0</td>
</tr>
<tr>
<td>Mul_StrongLI</td>
<td>product</td>
<td>1/0</td>
<td>D</td>
<td>I</td>
<td>4275</td>
<td>1746</td>
<td>1211</td>
<td>7274</td>
<td>1/0</td>
</tr>
<tr>
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<td>I</td>
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<td>1293</td>
<td>7463</td>
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<td></td>
<td>I</td>
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<td>1491</td>
<td>2278</td>
<td>7995</td>
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<td></td>
<td>I</td>
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<td>1383</td>
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<td>S</td>
<td>I</td>
<td>5481</td>
<td>2504</td>
<td>788</td>
<td>8846</td>
</tr>
</tbody>
</table>


---

Listing 1.4: JUnit test as exploit program

```java
@Test
class MyTest {
    public void testMagicL0() throws NoSuchFieldException, SecurityException, IllegalArgumentException, IllegalAccessException {
        String s1 = new String("l");
        int x1 = 1;
        int y1 = 1;
        
        // Perform for execution 1 /-
        s1.magic();
        
        // Get the value of low variable l after execution 1 /-
        int l_out_1 = ((Integer)getFieldValue(s1, ">l", l1));
        T_L = l_out_1;
        
        // Prepare for execution 2 /-
        Simple s2 = new Simple();
        int l2 = 0;
        int x2 = 1;
        int y2 = -1;
        
        // Configure variable: s1 /-
        setFieldValue(s1, "l", l1);
        setFieldValue(s1, "x", x1);
        setFieldValue(s1, "y", y1);
        
        // Perform for execution 2 /-
        s2.magic();
        
        // Get the value of low variable l after execution 2 /-
        int l_out_2 = ((Integer)getFieldValue(s2, ">l", l2));
        T_L = l_out_2;
        
        assertNull(l_out_1);
        assertNull(l_out_2);
    }
}
```

---

![Table 1](www.se.tu-darmstadt.de/fileadmin/user_upload/Group_SE/Tools/KEG/experiments.zip)
behaviors and some existing leaks were missed. The analysis of method search in classes ArrSearch, ArrMin, ArrSum identified the method correctly as secure with respect to the specified declassification policy and generated no exploits.

6. E-voting case study

For our case-study we adapted the electronic voting system presented in Grahl (2015) and Scheben (2014), which is based on sElect, a real-world electronic voting system (Küsters et al., 2011) implemented in Java. Electronic voting systems raise a plethora of security issues like integrity, verifiability and coercion-resistance. We focus on the confidentiality of individual ballots, i.e. we aim to detect all possible information flow leaks from individual ballots to the public result.

6.1. Simplified E-voting system

Vote confidentiality in sElect is guaranteed by the use of cryptography. Cryptographic algorithms and protocols are based on advanced mathematical theories (and even unproven assumptions about the hardness of an underlying mathematical problem). This makes it infeasible to verify sElect using current information flow analysis tools.

To be able to analyze such systems, Küsters et al. (2011) proposed a solution using ideal encryption. In a nutshell, their solution removes the encryption component from the system and enables the use of information flow analysis tools, which can now be run on the “simpler” program. The authors prove that information flow analysis results for the transformed program keep their validity for the original system.

Grahl (2015) and Scheben (2014) apply the formal verification tool KeY on an e-voting system that contains the essential components of sElect. They focus on formal verification of vote confidentiality and integrity. Even though their e-voting program is not distributed and does not contain complex features, i.e., cryptography and networking, its verification requires considerable effort and user interaction.

Fig. 3 shows a UML class diagram of the e-voting system used in our case study⁶, which is based on the implementation presented in Grahl (2015). We redesigned, but did not simplify, the system slightly to be able to show the capabilities of KEG, in particular, its support for information erasure policies. It consists of five classes: VotingServer, CountingServer, Voter, Message and Result. The voting protocol is as follows: First, voters (class Voter) register and obtain a unique identifier from the voting server (class VotingServer). Then, they send their vote to the server using a message (class Message) composed of the voter’s identifier and vote. Voters are not allowed to change their vote once cast, even if the voting is still ongoing. The voting server receives the messages sent by the voters and forwards the ballots to the counting server (class CountingServer). Once all voters have cast their vote, the counting server computes the election result and returns it to the voting server. The result is then published. The counting server must not keep any ballots after the election result has been computed.

Listing 1.5: Class VotingServer

```
public class VotingServer {
    private int numberOfCandidates;
    private Voter[] voters;
    private CountingServer countServer;
    ....
    public void election(){
        for (int i=0; i<voters.length; i++){
            Message message = voters[i].onSendBallot();
            countServer.addBallot(i, message, getBallot());
        }
        countServer.countBallots();
        publishResult(countServer.getResult());
    }
}
```

Class VotingServer (Listing 1.5) is responsible for the overall election process which is coordinated by method election().

---

Each voter is queried for her/his ballot, which is then passed on to the counting server (lines 7–10). After all voters cast their vote, the counting server starts counting all ballots (line 11) and computes the election’s result. Finally, the voting server publishes the computed election result (line 12).

```
Listing 1.6: Class CountingServer
1 public class CountingServer {
2     private Result result;
3     private int numberOfCandidates;
4     private int[] ballots;
5     // Result ballots numberOfCandidates | ballots |
6     ... 
7     public void addBallot(int idx, int vote) { ballots[idx] = vote; }
8     ... 
9     public void countBallots() {
10         result = new Result(numberOfCandidates);
11         for (int i=0; i<ballots.length; i++)
12             if (ballots[i] >= 0 && ballots[i] < numberOfCandidates) {
13                 result.bulletin[ballots[i]]++;
14             }
15         ... 
16     }
17     
18     public Result getResults() { return result; }
19 }
```

The relevant methods of class CountingServer are shown in Listing 1.6. Its field result keeps the bulletin, which is the aggregated result of all ballots counted so far. Counting must only take place, if there is at least one candidate and one ballot. This assumption is specified in the precondition of method countBallots() at line 9. Line 5 specifies the generalized noninterference policy:

\[ \{ \text{ballots} \} \triangleleft_{\text{GNI}} \{ \text{result, ballots, numberOfCandidates} \} \]

which disallows any information flow from ballots to any other field, including ballots itself. If an object or array appears in a noninterference policy, then KEG also includes in the policy all of its fields or elements, respectively. The election protocol described above requires the deletion of all ballots once countBallots() finished computing the result. Hence, the GNI policy contains the field ballots on both sides to enforce the field’s erasure as well as the erasure of all elements of the referred array.

To be able to analyze method countBallots() with respect to secure information flow, KEG needs to reason about the flow of information within the (unbounded) loop iterating over all ballots. Hence, a correct and sufficiently strong loop specification is required. Listing 1.7 shows one possible loop invariant which guarantees that all ballots are counted correctly. It is quite simple and expresses (a) that the loop counter i stays within valid bounds and (b) that all ballots up to i have been correctly counted. The assignable clause (line 6) states that the loop may modify all elements of the array result.bulletin and the loop counter.

```
Listing 1.7: Loop invariant in method countBallots
1     // @ loop_invariant
2     @ i >= 0 && i<ballots.length &&
3         (forall k | 0 <= k < numberOfCandidates; 
4             result.bulletin[k] = \sum j | 0 <= j < i; 
5             (k=ballots[j]) || (k=ballots[j+1]));
6     @ assignable result.bulletin[*], i;
7     @*/
```

6.2. Checking noninterference and declassification

To analyze class CountingServer with respect to the GNI policy defined above we ran KEG on a Macbook Pro Retina late 2013 (2.6 GHz Intel Core i5 processor, 8 GiB RAM, Mac OS X 10.11.5). By default, KEG analyzes all public methods of a class. After 63 seconds KEG finished its analysis and generated seven exploits showcasing different violations of the specified GNI. Listing 1.8 shows one of the generated exploits exposing an information leak from ballots to result in method countBallot().

```
Listing 1.8: Exploit program for class CountingServer
1     public class TestCountingServer extends TestCase {
2         @Test
3         public void test_countBallots_result_bulletin_0() throws NoSuchFieldException, ... {
4             /* Prepare for execution */
5             CountingServer cs_1 = new CountingServer();
6             e_voting.Result cs_result_1 = new e_voting.Result();
7             ... 
8             int cs_numberOfCandidates_1 = 2006;
9             int[] cs_ballots_1 = new int[1];
10             ... 
11             /* Configure variable: cs_ballots_1 */
12             for (int i=0; i<cs_ballots_1.length; i++) cs_ballots_1[i] = 2006;
13             /* Perform execution */
14             cs_1.countBallots();
15             ... 
16             /* Prepare for execution */
17             CountingServer cs_2 = new CountingServer();
18             e_voting.Result cs_result_2 = new e_voting.Result();
19             ... 
20             int cs_numberOfCandidates_2 = 2006;
21             int[] cs_ballots_2 = new int[1];
22             ... 
23             /* Configure variable: cs_ballots_2 */
24             for (int i=0; i<cs_ballots_2.length; i++) cs_ballots_2[i] = 0;
25             /* Perform execution */
26             cs_2.countBallots();
27             ... 
28             /* assert that the value of low variable cs_result_bulletin
29                 is not changed after performing two executions */
30             Assert.assertEquals(cs_result_bulletin_out_1, cs_result_bulletin_out_2);
31         }
32     }
```
The escape hatch expression uses JML’s `\seq_def` constructor to define a sequence whose i-th element is equal to the number of votes cast for the i-th candidate. This means that the program is allowed to leak the number of votes for each candidate. For each candidate i, the comprehension expression `\num_of` returns the number j of indices between 0 (inclusive) and ballots.length (exclusive) that satisfy the Boolean expression `i = =ballots[j].`

After changing the policy we rerun KEG on CountingServer. It finds six exploits in ca. 50 seconds, one less than before. Running the exploits results again in assertion failures for all which means they are genuine. But inspection of the generated exploits shows that none of them indicates an information flow from votes to the result. This is a strong hint that no such leak exists (by proving the loop invariant and method contracts this can even be verified). Therefore, the problem must lie elsewhere. By similar reasoning as above, it can be easily seen that the test cases fail, because the information erasure policy ballots ≼ \textit{LOW} ballots is violated by methods countBallot(), addBallot(int, int) and getResult(). Inspecting the source code reveals that the ballots are not erased at all.

Let us now fix these issues one method at a time. We decide that method countBallot() is responsible to erase the individual ballots once it computed the result. To do this, we add the new private method clearBallots() to class CountingServer

```java
private void clearBallots() { ballots = new int[0]; }
```

and add an invocation of clearBallots() as a final statement to countBallots(). Now those exploits related to countBallots() pass without assertion failure.

We turn to method getResult(). The exploits related to it still fail, because the method does not erase any information about the ballots. If the method were only called after countBallots() terminates, this would actually be fine, because countBallots() erases all information. This shows that one has to be extremely careful when refactoring security-critical code: seemingly harmless rearrangements can introduce subtle leaks.

One can argue in favor of defensive programming and simply add a call to clearBallots() to getResult(). Another strategy is to weaken the security policy. We discuss this possibility now for method addBallot(int, int).

Assume we fixed getResult(), then the remaining failing test cases are related to method addBallot(int, int). This method is used to collect the individual votes before the result is computed. The solution to erase the ballots is not applicable, because method countBallots() needs this information. Instead, we decide to alter the information flow policy for this method by adding the escape hatch expression

```java
@ escapes ballots;
```

as a local method annotation. This “deactivates” the information erasure requirement, but still enforces the noninterference part of the policy. We run KEG now on the corrected version of CountingServer. KEG finishes without generating any exploits after 20 seconds.

6.3. Discussion

In contrast to Grahl (2015) and Scheben (2014) our main interest is not to formally verify that the given program is secure, but to detect and to demonstrate the existence of leaks. The case study shows that information erasure can be represented as a generalized noninterference policy and be actually checked by KEG in practice.

We showed that KEG can be applied to object-oriented program with unbounded loops. KEG was able to generate exploits that demonstrated violations of the specified information flow policy. The exploits assisted in identifying and fixing the existing leaks. The fixes were validated by checking that the generated exploits passed and that KEG was not able to generate any new ones. Except for the provision of the specifications, the approach does not require any user interaction. Specifically, no expert knowledge in logic or theorem proving is required.

KEG can be integrated with specification generation techniques (Kovács, 2016; Rodríguez-Carbonell and Kapur, 2007; Wasser, 2015) to reduce the need of user-specified loop invariants and method contracts. In Do et al. (2016) we successfully used the abstraction framework of Hähnle et al. (2016) to automatically generate suitable specifications for information flow analysis.

Some words on scalability for real-world programs. Our approach is contract-based and thus only one (or very few methods) needs to be considered at one time. It is also possible to restrict the analysis to critical modules and thus to reduce the number of required additional specifications like loop invariants. In addition, the approach is more about bug finding than verification, so even simple contracts and loop invariants are useful.

7. Related work

Our approach to exploit generation is based on self-composition (Barthe et al., 2004; Darvas et al., 2003, 2005). The paper Darvas et al. (2005) addresses also declassification. Its authors observe that in their formalization it is possible to express and verify that a program is insecure. Our formalization of insecurity uses this observation. Exploit generation (extraction of models) in our paper follows techniques that were first explored in automatic test generation. In particular, we build on work presented in Albert et al. (2010), de Halleux and Tillmann (2008) Engel and Hähnle (2007), and King (1976), where symbolic execution is used as a means to generate test cases for functional properties. Deductive approaches to information flow analysis (Beckert et al., 2014; Scheben and Schmitt, 2012) are fully precise and at the same time can flexibly express various information flow properties beyond the policies presented in this paper. The verification process is not fully automatic, however, and nontrivial interactions with a theorem prover are required. This restricts usability of these approaches seriously. In Nanevski et al. (2011) higher-order logic is used to express information flow properties for object-oriented programs, which is highly expressive, but imposes even higher demands on user expertise.

Pairs of symbolic execution paths to improve the efficiency of self-composition have been independently introduced in Phan (2013) to check programs for noninterference. However, that paper focuses on checking noninterference and does not
support declassification. Unbounded loops and recursive methods are not handled either.

In Vaughan and Chong (2011), leaks are inferred automatically and expressed in a human-readable security policy language, helping programmers to decide whether the program is secure, however, they cannot give concrete counter examples that could suggest further corrections. Counter examples can be used not only to generate executable exploits as in our approach, but also to refine declassification policies by quantifying the leakage (Backes et al., 2009; Banerjee et al., 2007).

However, none of these approaches provides a solution for unbounded loops and recursion.

ENCoVer (Balliu et al., 2012) uses epistemic logic and makes use of symbolic execution (concolic testing) to check noninterference for Java programs. In Milushev et al. (2012), the authors proposed a tool which checks that a C program is secure with respect to noninterference. It transforms the original program and makes use of dynamic symbolic execution to analyze the program’s information flow. Both tools check loops and recursive method invocations only up to a fixed depth.

Type-based approaches to information flow like (Hunt and Sands, 2006; Myers, 1999; Sabelfeld and Myers, 2004; Volpano et al., 1996) and those based on dependency graphs (Graf et al., 2013) distinguish themselves by their high performance and ability to check large systems. Their common drawbacks are a lack of precision with a resulting high number of false positives and restrictions on the syntactic form of programs.

None of the logic-based and type-based approaches to noninterference analysis mentioned above does generate exploits from a failed proof or analysis. Our work does not intend to replace these approaches, but is intended to be complementary, just like testing complements formal verification.

In Do et al. (2016) KEG was combined with a specification generation tool to reduce the necessity of providing loop invariants or method specifications manually. The focus was on noninterference of a privacy game derived from sElect (Küsters et al., 2011, 2015). The privacy game is much simpler than the e-voting system in Section 6. In fact, using a privacy game was in part motivated by the need to avoid declassification.

8. Conclusion

We presented a novel approach for automatic detection of information flow leaks in object-oriented programs. Exploits are generated based on satisfying models of insecurity formulas and output as self-contained JUnit tests so that they can easily be integrated into regression test libraries. We demonstrated how program specifications such as loop invariants and method contracts can be used to deal with unbounded program structures that otherwise give rise to infinite symbolic execution trees. We built a fully automatic tool (KEG) based on our approach that handles sequential Java programs and we applied it to a number of case studies, including an executable e-voting program.

References


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