Enforcing Information Flow Policies with Type-Targeted Program Synthesis

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We present a technique for static enforcement of high-level, declarative information flow policies. Given a program that manipulates sensitive data and a set of declarative policies on the data, our technique automatically inserts policy-enforcing code throughout the program to make it provably secure with respect to the policies. We achieve this through a new approach we call type-targeted program synthesis, which enables the application of traditional synthesis techniques in the context of global policy enforcement. The key insight is that, given an appropriate encoding of policy compliance in a type system, we can use type inference to decompose a global policy enforcement problem into a series of small, local program synthesis problems that can be solved independently.

We implement this approach in Lifty, a core DSL for data-centric applications. Our experience using the DSL to implement three case studies shows that (1) Lifty’s centralized, declarative policy definitions make it easier to write secure data-centric applications, and (2) the Lifty compiler is able to efficiently synthesize all necessary policy-enforcing code, including the code required to prevent several reported real-world information leaks.

1 INTRODUCTION

From social networks to health record systems, today’s software manipulates sensitive data in increasingly complex ways. To prevent this data from leaking to unauthorized users, programmers sprinkle policy-enforcing code throughout the system, whose purpose is to hide, mask, or scramble sensitive data depending on the identity of the user or the state of the system. Writing this code is notoriously tedious and error-prone.

Static information flow control techniques [Chlipala 2010; Jia and Zdancewic 2009; Li and Zdancewic 2005; Myers 1999; Swamy et al. 2010; Zheng and Myers 2007] mitigate this problem by allowing the programmer to state a high-level declarative policy, and statically verify the code against this policy. These techniques, however, only address part of the problem: they can check whether the code as written leaks information, but they do not help programmers write leak-free programs in the first place. In this work, we are interested in alleviating the programmer burden associated with writing policy-enforcing code.

In recent years, program synthesis has emerged as a powerful technology for automating tedious programming tasks [Barowy et al. 2015; Feng et al. 2017; Gulwani 2011; Solar-Lezama et al. 2006; Yaghmazadeh et al. 2017]. In this paper we explore the possibility of using this technology to enforce information flow security by construction: using a declarative policy as a specification, our goal is to automatically inject provably sufficient policy-enforcing code throughout the system.
This approach seems especially promising, since each individual policy-enforcing snippet is usually short, side-stepping the scalability issues of program synthesizers.

The challenge, however, is that our setting is significantly different from that of traditional program synthesis. Existing synthesis techniques target the generation of self-contained functions from end-to-end specifications of their input-output behavior. In contrast, we are given a global specification of one aspect of the program behavior: it must not leak information. This specification says nothing about where to place the policy-enforcing snippets, let alone what each snippet is supposed to do.

**Type-targeted program synthesis.** In this paper we demonstrate how to bridge the gap between global policies and local enforcement via a new approach that we call type-targeted program synthesis. Our main insight is that a carefully designed information flow type system lets us leverage type error information to infer local, end-to-end specifications for policy-enforcing leak patches. More specifically, (1) the location of a type error indicates where to insert a patch and (2) its expected type corresponds to the local specification for the patch. A crucial property of local specifications is that any combination of patches that satisfy their respective local specifications yields a provably secure program. In other words, type-targeted synthesis decomposes the problem of policy enforcement into several independent program synthesis problems, which can then be tackled by state-of-the-art synthesis techniques.

**Type system.** The main technical challenge in making type-targeted synthesis work is the design of a type system that, on the one hand, is expressive enough to reason about the policies of interest, and on the other hand, produces appropriate type errors for the kinds of patches we want to synthesize. For our policy language, we draw inspiration from Jeeves [Austin et al. 2013; Yang et al. 2012], which supports rich, context-dependent policies, where the visibility of data might depend both on the identity of the viewer and the state of the system. For example, in a social network application, a user’s birth date can be designated visible only to their friends (where the list of friends can be updated over time). In Jeeves, these policies are expressed directly as predicates over users and states. Our technical insight is that static reasoning about Jeeves-style policies can be encoded in a decidable refinement type system by indexing types with policy predicates. Moreover, we show how to instantiate the Liquid Types framework [Rondon et al. 2008] to infer appropriate expected types at the error locations.
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The **LIFTY language.** Based on this insight, we developed LIFTY\(^1\), a core DSL for writing secure data-centric applications. In LIFTY, the programmer implements the core functionality of an application as if all data were publicly visible. Separately, they provide a policy module, which associates declarative policies with some of the fields (columns) in the data store, by annotating their types with policy predicates. Given the source program and the declarative policies, LIFTY automatically inserts leak patches across the program, so that the resulting code provably adheres to the policies (Fig. 1). To that end, LIFTY’s type inference engine checks the source program against the annotated types from the policy module, flagging every unsafe access to sensitive data as a type error. Moreover, for every unsafe access the engine infers the most restrictive policy that would make this access safe. Based on this policy, LIFTY creates a local specification for the leak patch, and then uses a variant of type-driven synthesis [Polikarpova et al. 2016] to generate the patch.

**Evaluation.** To demonstrate the practical promise of our approach, we implemented a prototype LIFTY-to-Haskell compiler. We evaluated our implementation on a series of small but challenging micro-benchmarks, as well as three case studies: a conference manager, a health portal, and a student grade record system. The evaluation demonstrates that our solution supports expressive policies, reduces the burden placed on the programmer, and is able to generate all necessary patches for our benchmarks within a reasonable time (26 seconds for our largest case study). Importantly, the evaluation confirms that the patch synthesis time scales linearly with the size of the source code (more precisely, with the number of required leak patches), suggesting the feasibility of applying this technique to real-world code bases.

2 LIFTY BY EXAMPLE

To introduce LIFTY’s programming model, type system, and the type-targeted synthesis mechanism, we use an example based on a leak from the EDAS conference manager [Agrawal and Bonakdarpour 2016]. We have distilled our running example to a bare minimum to simplify the exposition of how LIFTY works under the hood; at the end of the section, we demonstrate the flexibility of our language through more advanced examples.
showPaper client p = let row = do t ← get (title p) st ← get (status p) ses ← if st = Accept then get (session p) else "" t + " " + ses in print client row

Fig. 3. Snippet of the core functionality of a conference manager (in Lifty syntax).

showPaper client p = let row = do t ← get (title p) st ← do x1 ← get phase if x1 = Done then get (status p) else NoDecision ses ← if st = Accept then get (session p) else "" t + " " + ses in print client row

Fig. 4. With a leak patch inserted by Lifty.

2.1 The EDAS Leak

Fig. 2 shows a screenshot from the EDAS conference manager. On this screen, a user can see an overview of all their papers submitted to upcoming conferences. Color coding indicates paper status: green papers have been accepted, orange have been rejected, and yellow papers are awaiting notification. The user is not supposed to learn the acceptance decision before the notifications are out, yet from this screen they can infer that the first one of the yellow papers has been tentatively accepted, while the second one has been tentatively rejected. They can make this conclusion because the two rows differ in the value of the “Session” column (which displays the conference session where the paper is to be presented), and the user knows that sessions are only displayed for accepted papers.

The EDAS leak is particularly insidious because it provides an example of an implicit flow: the “accepted” status does not appear anywhere on the screen, but rather influences the output via a conditional. To prevent such leaks, it is insufficient to simply examine output values; rather, we must track the flow of sensitive information through the system.

Fig. 3 shows a simplified version of the code that has caused this leak. This code retrieves the title and status for a paper p, then retrieves session only if the paper has been accepted, and finally displays the title and the session to the currently logged-in client. The leak happened because the programmer forgot to insert policy-enforcing code that would prevent the true value of status from influencing the output, unless the conference is in the appropriate phase (notifications are out). It is easy to imagine, how in an application that manipulates a lot of sensitive data, such policy-enforcing code become ubiquitous, imposing a significant burden on the programmer and obscuring the application logic.

2.2 Programming with Lifty

Lifty liberates the programmer from having to worry about policy-enforcing code. Instead, they provide a separate policy module that describes the data layout and associates sensitive data with declarative policies. For example, Fig. 5 shows a policy module for our running example.

The Lifty type system is equipped with a special type constructor $⟨T⟩^{\pi}$ ("$T$ tagged with policy $\pi$"), where $\pi : (\Sigma, \text{User}) \rightarrow \text{Bool}$ is a predicate on output contexts, i.e. pairs of stores and users. The type $⟨T⟩^{\lambda(s,u),p}$ denotes values of type $T$ that are only visible to a user $u$ in a store $s$ such that $p$ holds.

1Lifty stands for Liquid Information Flow TTypes.
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module ConfPolicy where

data Status = Accept | Reject | NoDecision

redact NoDecision

title :: PaperId → Ref (String)^λ T

status :: PaperId → Ref (Status)^λ(s,u).s[phase] = Done

session :: PaperId → Ref (String)^λ T

Fig. 5. Snippet from a policy module for a conference manager.

For example, to express that a paper’s status is only visible when the conference phase is Done, the programmer defines its type as a reference to Status tagged with policy λ(s,u).s[phase] = Done. Hereafter, we elide the binders (s, u) from policy predicates for brevity, and simply write λ.p. The policy λ.T annotates fields as public (i.e. visible in any context).

Given the code in Fig. 3 and the policy module, LIFTY injects policy-enforcing code required to patch the EDAS leak; the result is shown in Fig. 4 with the new code highlighted. This code guards the access to the sensitive field status with a policy check, and if the check fails, it substitutes the true value of status with a redacted value—a constant NoDecision. LIFTY constructs redacted values using a restricted set of functions, designated by the programmer via redact clauses in the policy module (Fig. 5). LIFTY guarantees that the patched code is provably correct with respect to the policies in the policy module.

Even though in this example generating the policy check amounts to little more than copying the policy predicate from the type declaration, this naive syntactic approach quickly breaks down in more realistic scenarios. In particular, the policy might depend on the eventual viewer, the state of the system might have changed in between the invocations of get and print, or there might be multiple choices for constructing the redacted value. To be able to handle these scenarios, we opt for a more general approach based on program synthesis.

2.3 Type-Targeted Program Synthesis

Can the code in Fig. 4—and its correctness proof—be synthesized using existing techniques? Several existing synthesis systems [Kneuss et al. 2013; Polikarpova et al. 2016] provide correctness guarantees, but they generate programs from scratch, from full functional specifications. In our case, a full functional specification for showPaper is not available; in addition, re-synthesizing the whole function from scratch will fail to scale to larger programs.

Prior approaches to sound program repair [Kneuss et al. 2015] use fault localization to focus synthesis on small portions of the program responsible for the erroneous behavior. These existing localization techniques, however, are insufficient in our setting. First, they are not precise enough, i.e. they would fail to pinpoint get (status p) as the smallest unsafe term; to increase precision, they rely on testing, which is hardly applicable to non-functional properties such as information flow security. Second, they can only identify where to insert the patch, but cannot infer its local specification; candidate patches are validated by re-checking the whole function, which leads to a combinatorial explosion in the presence of multiple unsafe terms (a rare case in the context of program repair, but quite common with the LIFTY programming model).
In this section we show how a careful encoding of information flow security into a type system (Sec. 2.3.1) allows us to instead use type inference for precise fault localization (Sec. 2.3.2). Concretely, type-checking the code in Fig. 3 against the policy module, leads to a type error in line 5, which flags the term \texttt{get (status p)} as unsafe, and moreover, gives its expected type, which can be used as the local specification for patch synthesis (Sec. 2.3.3).

2.3.1 Type System. The Lifty type system builds upon existing work on security monads [Russo et al. 2008; Swamy et al. 2010], where sensitive data lives inside a monadic type constructor (in our case, \langle \cdot \rangle), parameterized by a security level; proper propagation of levels through the program is ensured by the type of the monadic \texttt{bind} \footnote{Lifty’s Haskell-like do-notation used in Fig. 3 desugars into invocations of \texttt{bind} in a standard way [Marlow 2010] (see Appendix A for the desugared version).}. In contrast with prior work, our security levels correspond directly to policy predicates, which allows Lifty programs to express complex context-dependent policies directly as types, instead of encoding them into an artificial security lattice.

Subtyping. Moreover, unlike prior work, Lifty features subtyping between tagged types, which is contravariant in the policy predicate, i.e. \langle T \rangle^\alpha \pi < \langle T \rangle^\beta \pi \iff \pi \Rightarrow p. This allows a “low” value (with a less restrictive policy) to appear where a “high” value (with more restrictive policy) is expected, but not the other way around. Lifty restricts the language of policy predicates to decidable logics; hence, the subtyping between tagged types can be automatically decided by an SMT solver.

Tagged primitives. The type error for the EDAS leak is generated due to the typing rules for primitive operations on tagged values, \texttt{bind} and \texttt{print}. Informally, the \texttt{bind} rule allows tagging the result of a sequence of two tagged computations with any policy \pi that is as least as secret as both computations. The \texttt{print} rule allows displaying messages tagged with any \pi that holds of the current state and the viewer. We formalize these rules in Sec. 3.

Type inference. The Lifty type inference engine is based on the Liquid Types framework [Cosman and Jhala 2017; Rondon et al. 2008]. As such, it uses the typing rules to generate a system of subtyping constraints over tagged types, and then uses the definition of contravariant subtyping to reduce them to the following system of Horn constraints over policy predicates:

\begin{align*}
B_4 & \Rightarrow s[\text{phase}] = \text{Done} \quad (1) \\
B_3 & \Rightarrow B_4 \quad (2) \\
u = \text{client} \land s = \sigma & \Rightarrow B_3 \quad (3)
\end{align*}

where \(B_i\) are unknown policies of \texttt{bind} applications at line \(i\) (trivial constraints are omitted). Specifically, (1) and (2) collectively state that the \texttt{row} computation must be at least as secret as the status field, while (3) requires \texttt{row}’s policy to hold of the output context.

The Horn constraints are solved using a combination of unfolding and predicate abstraction. In this case, however, the system clearly has no solution, since the consequent of (1) is not implied by the antecedent of (3); in other words, the code is trying to display a sensitive value in a context where its policy doesn’t hold.

2.3.2 Fault Localization. Unlike existing refinement type checkers [Cosman and Jhala 2017; Rondon et al. 2008], Lifty is not content with finding that a type error is present: it needs to identify the term to blame and infer its expected type. Intuitively, declaring constraint (3) as the root cause of the error corresponds to blaming \texttt{print} for displaying its sensitive message in too many contexts, while picking constraint (1), corresponds to blaming the access \texttt{get (status p)} for returning a value that is too sensitive. For reasons explained shortly, Lifty always blames the access.
To infer its expected type, it has to find an assignment to $B_4$, which works for the rest of the program (i.e., a solution to constraints (2)–(3)). This new system has multiple solutions, including a trivial one $[B_i \mapsto \top]$. The optimal solution corresponds to the least restrictive expected type, in other words—due to contravariance—the strongest solution for policy predicates: $[B_i \mapsto u = \text{client} \land s = \sigma]$. Substituting this solution into the subtyping constraint that produced (1), results in the desired type error:

```
get (status p):
  expected type: $\langle \text{Status} \rangle \lambda \cdot u = \text{client} \land s = \sigma$
  and got: $\langle \text{Status} \rangle \lambda \cdot s[\text{phase}] = \text{Done}$
```

Note that picking constraint (3) as the root cause instead, and inferring the weakest solution to constraints (1)–(2) ($[B_i \mapsto s[\text{phase}] = \text{Done}]$) conceptually would give rise to a different patch: guarding the whole message row with a policy check. This would fix the leak but have an undesired side effect of hiding the paper title along with the session. Data-centric applications routinely combine multiple pieces of data with different policies in a single output; therefore, in this domain it makes more sense to guard the access, which results in “redacting” as little data as possible.

```
\text{2.3.3 Patch Synthesis.} From the expected type, LIFTY obtains a type-driven synthesis problem [Polikarpova et al. 2016]:

\[ \Gamma \vdash ?? :: \langle \text{Status} \rangle \lambda \cdot u = \text{client} \land s = \sigma \]

Here $\Gamma$ is a \textit{typing environment}, which contains a set of components—variables and functions that can appear in the patch—together with their refinement types. A solution to this problem is any program term $t$ that type-checks against the expected type in the environment $\Gamma$. As we show in Sec. 4, any such $t$, when substituted for \textit{get} (status p) in Fig. 3, would produce a provably secure program; hence the synthesis problem is local (can be solved in isolation).

Even though any solution is secure, not all solutions are equally desirable: for example, we wouldn’t want the patch to return Accept unconditionally. Intuitively, a desirable solution returns the original value whenever allowed, and otherwise replaces it with a reasonable redacted value. LIFTY achieves this through a combination of two measures. First, instead of synthesizing a single term of type $\langle \text{Status} \rangle \pi$, it generates a set of candidate branches (by enumerating all branch-free terms of this type up to a fixed size). Second, LIFTY gives the programmer control over the set of possible redacted values by generating the branches in a restricted environment, which only contains the original term and the components explicitly designated by the programmer as redaction functions (see Fig. 5 for an example). As a result, our running example generates only two branches:

```
get (status s):
  $\langle \text{Status} \rangle \lambda \cdot s[\text{phase}] = \text{Done}$
NoDecision:
  $\langle \text{Status} \rangle \lambda \cdot \top$
```

Next, for every branch, LIFTY attempts to abduce a condition that would make the branch type-check against the expected type. In our example, the second branch is correct unconditionally, while the first branch generates the following \textit{logical abduction} problem:

```
C \land u = \text{client} \land s = \sigma \Rightarrow s[\text{phase}] = \text{Done}
```

where $C$ is an unknown formula that cannot mention the policy parameters $s$ and $u$. LIFTY uses existing techniques [Polikarpova et al. 2016] to find the following solution

```
C \mapsto \sigma[\text{phase}] = \text{Done}.
```

\text{\textsuperscript{3}}LIFTY borrows this “redaction semantics” from Jeeves, which also enforces policies by replacing secret values in a computation with public defaults, but does so via a specialized runtime (see a detailed comparison in Sec. 6).
It then sorts all successfully abduced branch conditions from strongest to weakest, and uses each condition to synthesize a guard, i.e. a program that computes the monadic version of the condition:

\[
\text{bind } (\text{get phase}) \ (x_1 \ . \ x_1 = \text{Done}).
\]

Finally, LIFTY combines the synthesized guards and branches into a single conditional, which becomes the patch and replaces the original unsafe access.

### 2.4 Scaling Up to Real-World Policies

In the rest of the section, we demonstrate more challenging scenarios, where (1) a function contains several unsafe accesses, (2) the policy check itself uses sensitive data, and hence proper care must be taken to ensure that policy-enforcing code does not introduce new leaks, (3) the redacted value is not just a constant, or (4) the policy check depends on the eventual viewer and the state at the time of output (which need not equal the state at the time of data retrieval). Full code for each of these examples can be found in Appendix B.

#### 2.4.1 Multiple Leaks

Consider a variant of our running example, where in addition to the paper’s title and session, we display its authors. Also assume our conference is double-blind, so the authors field is only visible once notifications are out:

\[
\text{authors} :: p : \text{PaperId} \rightarrow \text{Ref} \langle \text{[User]} \rangle \lambda \ s \ [\text{phase}] = \text{Done}
\]

When checking this extended version of `showPaper`, LIFTY generates two type errors, one for `get(status p)` with expected type `⟨\text{Status}\rangle^\pi` and one for `get(authors p)` with expected type `⟨\text{[User]}\rangle^\pi`, where \(\pi \equiv \lambda \ u = \text{client} \land \ s = \sigma\) in both cases, since both faulty terms flow into the same `print` statement. This gives rise to two patch synthesis problems, which can be solved independently, because their expected types only depend on the output context.

Note that LIFTY is able to infer independent expected types for the two errors because the correctness of this program depends only on policies of the faulty terms and not on their content. This is a crucial property that enables local patch synthesis. We observe that, unlike in general program repair, this assumption is reasonable in the context of policy enforcement for data-centric applications (as confirmed by our evaluation).

#### 2.4.2 Self-Referential Policies

Continuing with our extended example, assume that we want to allow a paper’s author to see the author list even before the notifications are out. This is an example of a policy that depends on a sensitive value; moreover, in this case the policy is self-referential: it guards access to `authors` in a way that depends on the value of `authors`. Enforcing such policies by hand is particularly challenging, because the policy check needs to retrieve and compute over sensitive values, and hence, while trying to patch one leak, it might inadvertently introduce another.

In LIFTY, the programmer expresses this complex policy in a straightforward way:

\[
\text{authors} :: p : \text{PaperId} \rightarrow \text{Ref} \langle \text{[User]} \rangle \lambda \ s [\text{phase}] = \text{Done} \lor u \in s \ [\text{authors p}]
\]

Given this policy, LIFTY generates a provably correct patch:

\[
\text{auths} \gets \text{do } c_1 \gets \text{do } x_1 \gets \text{get phase}; \ x_1 = \text{Done} \\
\quad c_2 \gets [\text{do } x_2 \gets \text{get (authors p)}; \ \text{elem client} \ x_2] \\
\quad \text{if } c_1 \lor c_2 \text{ then get (authors p) else []}
\]

The policy check `c2` is wrapped in `\[\cdot\]`, LIFTY’s novel safe downgrading construct. This construct changes the tag on `c2` from “visible to authors p” to “visible to client”, which is necessary for the patched program to type-check. LIFTY deems this downgrading safe because it can prove that `c2` does not give client the power to distinguish between two author lists they are not allowed to
see (since for any such list \(c_2\) evaluates to \(false\)). In Sec. 3 we formalize a security guarantee that captures this kind of reasoning, and show how \textsc{Lifty} performs this reasoning automatically.

**2.4.3 Nontrivial Patches.** When sensitive data has more interesting structure, the optimal redacted value can be more complex than just a constant. Consider the example of an auction where bids are only revealed once all participants have bid [Russo et al. 2008]. Now consider a more interesting policy: once a participant has bid and before all bids are fully revealed, they can see who else has bid, but not how much. One way to encode this in \textsc{Lifty} is to store the bid in an option type and associate different policies with the option and its content:

\[
\text{bid} :: \text{User} \rightarrow \text{Ref}\langle \text{Maybe}\langle \text{Int}\rangle \rangle \lambda s[\text{phase}] = \text{Done} \\
\text{showBid client } p, \text{ which displays participant } p\text{'s bid to client:}
\]

1. \(\text{do } c_1 \leftarrow \text{do } x_1 \leftarrow \text{get phase}; \ x_1 = \text{Done}\)
2. \(\text{if } c_1\)
3. \(\text{then get (bid p)}\)
4. \(\text{else do } c_2 \leftarrow [\text{do } x_2 \leftarrow \text{get (bid client)}; \ \text{isJust } x_2]\)
5. \(\text{if } c_2\)
6. \(\text{then do } x_3 \leftarrow \text{get (bid p)}; \ \text{mbMap (λ } \_ \_ \_ \_ \_ \text{) } x_3\)
7. \(\text{else Nothing}\)

This patch has three branches, of which the second one (line 6) is the most interesting: whenever client has bid but the bidding is not yet done, \textsc{Lifty} only redacts the value that might potentially be stored inside \(x_3\), but not whether \(x_3\) is \text{Nothing} or \text{Just} (here \text{mbMap} is Haskell’s \text{fmap} specialized to \text{Maybe}). Note that \text{Lifty} proves safety of this branch solely based on the type of \text{mbMap}:

\[
\text{mbMap} :: (α \rightarrow β) \rightarrow \text{Maybe } α \rightarrow \text{Maybe } β
\]

**2.4.4 State Updates.** Continuing with the auction example, consider the implementation of the function \text{placeBid} client b, which first retrieves everyone’s current bids, then updates client’s bid, and finally displays all the previously retrieved bids to client. In this case, reusing the patch from above would result in hiding too much, since \(x_3\) would reflect client’s (missing) bid at the time of retrieval; by the time of output, however, client has already bid and has the right to see who else did. \text{Lifty} would insert a correct repair, since it can reason about state updates, and in this case can statically determine that \(s[\text{bid } u] \neq \emptyset\) holds of the output context.

### 3 \textsc{The \(\lambda^L\) Type System}

We now formalize the type system of a core security-typed language \(\lambda^L\), which underlies the design of \text{Lifty}. The main novelty of this type system is representing security labels as policy predicates. This brings two important benefits: on the one hand, our type system directly supports context-dependent policies; on the other hand, we show how to reduce type checking of \(\lambda^L\) problems to liquid type inference [Rondon et al. 2008]. As a result, our type system design enables automatic verification of information flow security against context-dependent policies, and requires no auxiliary user annotations. Moreover, Sec. 4 also demonstrates how this design enables precise fault localization required for type-targeted synthesis of policy-enforcing code.

Another novelty of the \(\lambda^L\) type system is its support for policies that depend on sensitive values, including self-referential policies (Sec. 2.4.2). Until now, these policies were only handled by runtime techniques such as Jeeves [Austin et al. 2013; Yang et al. 2012]. To support safe enforcement
of these policies, $\lambda^L$ includes a novel safe downgrading construct (Sec. 3.3), and features a custom security guarantee, which we call contextual noninterference (Sec. 3.4).

### 3.1 Syntax of $\lambda^L$

$\lambda^L$ is a simple core language with references, extended with several information-flow specific constructs. We summarize the syntax of $\lambda^L$ in Fig. 6.

**Program terms.** $\lambda^L$ differentiates between expressions and statements. Expressions include store read (get), monadic bind (bind), and safe downgrading ($\lfloor \cdot \rfloor$), which we describe in detail below. A statement can modify the store (set) or output a value to a user (print). In the interest of clarity, we restrict the presentation in this section to a version of $\lambda^L$ where each program contains a single set or print statement; we defer the presentation of the full language with multiple statement to Appendix C. Furthermore, in order to focus on the challenges introduced by context-dependent policies, we keep statements unconditional, and thus avoid the—tedious but standard—complications associated with implicit flows (such as having to keep track of the latent security level). Implicit flows have to be encoded by passing conditional expressions as arguments to set or print.

**Types.** $\lambda^L$ supports static information flow tracking via tagged types. The type $\langle T \rangle^\pi$ (“$T$ tagged with $\pi$”) attaches a policy predicate $\pi : (\Sigma, \text{User}) \rightarrow \text{Bool}$ to a type $T$ (here $\Sigma$ is the type of stores, which map locations to values). A tagged type is similar to a labeled type in existing security-typed languages [Pottier and Simonet 2003; Sabelfeld and Myers 2003; Swamy et al. 2010], except the domain of labels is not an abstract lattice, but rather the lattice of predicates over stores and users. Intuitively, a value of type $\langle T \rangle^{\lambda(s,u)\cdot p}$ can be revealed in any store $s$ to any user $u$, such that $p$ holds. Here $p$ is a refinement predicate over the program variables in scope and the policy parameters $s$ and $u$. The exact set of refinement predicates depends on the chosen refinement logic; the only requirement is that the logic be decidable to enable automatic type checking. We assume that the logic at least includes the theories of uninterpreted functions ($x$ and $r$ $r$) and arrays ($r[r]$ and $r[r := r]$), which $\lambda^L$ uses to encode policy predicates and store reads/writes, respectively.

Other types of $\lambda^L$ include primitive types, references, as well as reference types and dependent function types, which are standard [Knowles and Flanagan 2010; Rondon et al. 2008]. In a refined base types $\{B \mid r\}, r$ is a refinement predicate over the program variables and a special value variable $v$, which denotes the bound variable of the type.

**Constants.** As is standard in refinement types literature [Knowles and Flanagan 2010], the syntactic category of constants, $c$, include values of primitive types and built-in functions on them. To formalize the LIFTY’s notion of a policy module, we assume constants include a predefined set of store locations and fields (functions that return locations). The type of each constant $c$
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Expression Evaluation

\[
\begin{align*}
\sigma, e & \rightarrow e \\
\sigma, \text{get } l & \rightarrow \sigma[l] \\
\sigma, \text{downgrade } \lfloor v \rfloor & \rightarrow v \\
\sigma, \text{bind } v & \rightarrow [\lambda x : T. e]v \\
\end{align*}
\]

Statement Execution

\[
\begin{align*}
\sigma, \text{set } l v & \rightarrow \sigma[l := v] \\
\sigma, \text{print } u v & \rightarrow \sigma[u += v] \\
\end{align*}
\]

Fig. 7. λL operational semantics.

is determined by an auxiliary function \(\text{ty}(c)\). For example, in a conference manager we define \(\text{ty}(\text{title}) = \text{PaperId} \rightarrow \langle \text{String} \rangle\). Since λL programs do not allocate new references at run time, the type of any location \(l\) is known a-priori and can be obtained through \(\text{ty}(l)\), which is why our typing rules do not keep track of a "store environment".

3.2 Dynamic Semantics of λL

The run-time behavior of λL programs is straightforward; we present non-standard rules in Fig. 7 (the full set of rules can be found in Appendix C). Expression evaluation happens in the context of a store \(\sigma : (\text{Loc} \rightarrow \text{Value}) \cup (\text{User} \rightarrow \text{Value}^*)\), which has two disjoint components. The first component is standard and is read and modified through get and set, respectively; the second component represents each user’s output channel is modified through print.

The dynamic semantics of tagged primitives is trivial, since λL only tracks policies statically. \(\lfloor \cdot \rfloor\) is an identity function, while bind corresponds to the bind of the identity monad.

3.3 Static Semantics of λL

Fig. 8 shows a subset of subtyping and type checking rules for λL that are relevant to information flow tracking. Other rules are standard for languages with decidable refinement types [Rondon et al. 2008; Vazou et al. 2013, 2014a,b] and deferred to Appendix C. In Fig. 8, a typing environment \(\Gamma ::= e | \Gamma, x : T | \Gamma, r\) maps variables to types and records path conditions \(r\).

Subtyping. We only show subtyping rules for tagged types. The rule \(\lll\text{Tag1}\) allows embedding a pure computation into a public computation (it serves as an implicit monadic return). The rule \(\lll\text{Tag2}\) specifies that tagged types are contravariant in their policy parameter; this relation allows “upgrading” a term with a less restrictive policy (more public) into one with a more restrictive policy (more secret) and not the other way around. The premise \(\Gamma \models r' \Rightarrow r\) checks implication between the policies under the assumptions stored in the environment (which include path conditions and refinements on program variables). By restricting refinement predicates to decidable logic, we make sure that this premise can be validated by an SMT solver. To our knowledge, λL is the first security-typed language that supports automatic upgrading for expressive policies.

Term typing. The rest of Fig. 8 defines the typing judgments for expressions (\(\Gamma; \sigma \vdash e :: T\)) and statements (\(\Gamma; \sigma \vdash s\)). Since λL is stateful, the judgments keep track of the current store \(\sigma\), which is used in the rule T-get to precisely characterize the retrieved value. The rule for conditionals (P-If) is standard, but we include it because verification of programs with policy checks relies crucially on its path-sensitivity. For example, in Fig. 4, \(x_1\) is assigned the precise type \(\{\text{Phase} | v = \sigma[\text{phase}]\}\), and hence the sensitive term get (status p) is type-checked in an environment extended with
whenever we can prove that $e$ binds $\lambda$.

To this end, $e$ is both amenable to automatic verification and particularly useful for self-referential policies.

Hard to check automatically in general, but the special case where $e$ binds $\lambda$ is sufficient to prove that it complies with the expected policy.

The core of information flow checking in $\lambda_1$ are the rules $\text{T-BIND}$ and $\text{T-PRINT}$, which, in combination with contravariant subtyping, guarantee that tagged values only flow into allowed contexts. To this end, $\text{T-BIND}$ postulates that applying a sensitive function to a sensitive value, yields a result that is at least as secret as either of them. According to $\text{T-PRINT}$, a $\text{print}$ statement takes as input a tagged user (which may be computed from sensitive data) and a tagged result. The rule requires both arguments to be tagged with the same policy $\pi$, and crucially, $\pi$ must hold of the viewer in the current store (i.e. both the viewer identity and the message must be visible to the viewer). Here $\pi((\sigma, v))$ stands for “applying” the policy predicate; formally $(\lambda(s, u), p)((\sigma, v)) \doteq p[s \mapsto \sigma, u \mapsto v]$. 

**Downgrading.** The safe downgrading construct, $[e]$, is a novel feature of $\lambda_1$, which we introduced specifically to support static verification of programs with self-referential policies (Sec. 2.4.2). Informally, the idea is that we can we can safely downgrade a tagged term $e$ (i.e. weaken its policy) whenever we can prove that $e$ is constant, since constants cannot leak information. Constancy is hard to check automatically in general, but the special case where $e$ is a tagged boolean turns out to be both amenable to automatic verification and particularly useful for self-referential policies.

The rule $\text{T-\cdot\cdot}$ allows tagging $[e]$ with $\lambda(s, u), p$ as long as there exists a refinement term $r$ over program variables, such that $e$ can be tagged with $\lambda(s, u), p \land r$ and the value of $e$ implies $r$. This

Subtyping

\[
\begin{align*}
\Gamma \vdash T \ll T' &\quad \text{Expression Typing} \quad \Gamma;\sigma \vdash e :: T \\
\Gamma \vdash T \ll (T)_{\lambda, \cdot} &\quad \text{Expression Typing} \quad \Gamma;\sigma \vdash e :: T
\end{align*}
\]

Expression Typing

\[
\begin{align*}
\text{T-GET} &\quad \Gamma;\sigma \vdash x :: \text{Ref}\{B | r \land v = \sigma[x]\} \\
&\quad \Gamma;\sigma \vdash \text{get} x :: \text{Ref}\{B | r \land v = \sigma[x]\}
\end{align*}
\]

Statement Typing

\[
\begin{align*}
\text{T-PRINT} &\quad \Gamma;\sigma \vdash x_1 :: \langle\{\text{User} \mid \pi((\sigma, v))\}\rangle^{\pi} \\
&\quad \Gamma;\sigma \vdash \text{print} x_1 \ x_2
\end{align*}
\]

\[
\begin{align*}
\text{T-SET} &\quad \Gamma;\sigma \vdash x_1 :: \text{Ref} \quad \Gamma;\sigma \vdash x_2 :: T \\
&\quad \Gamma;\sigma \vdash \text{set} x_1 \ x_2
\end{align*}
\]

Fig. 8. Typing rules of $\lambda_1$. 

the path condition $T = (\sigma[\text{phase}] = \text{Done})$, which is sufficient to prove that it complies with the expected policy.
Enforcing Information Flow Policies with Type-Targeted Program Synthesis

operation is safe because whenever \( r \) holds, the two policies are the same; while whenever \( \neg r \) holds, the value of \( e \) is guaranteed to be \( \text{false} \) (a constant).

To illustrate the application of this rule, consider the following simple \( \lambda^L \) program:

```lambda
let x = [bind (get (authors p)) (λ ys . elem client ys)] in
print client x
```

where authors has the policy \( \lambda . u \in s[\text{authors } p] \). Even though client might not be an author, this program type-checks successfully, by instantiating the rule \( T - \lfloor \cdot \rfloor \) with \( \pi \equiv \lambda . u = \text{client} \land s = \sigma \) and \( r \equiv \text{client} \in \sigma[\text{authors } p] \). This program also satisfies contextual noninterference (Sec. 3.4): it does not allow client to distinguish between two author lists they shouldn’t see.

As a counter-example, replacing \( \text{elem client } x \) above with \( x = [] \) leads to a type error, since in this case \( x \) cannot be safely downgraded to \( \pi \) (no \( r \) satisfies the premise of \( T - \lfloor \cdot \rfloor \)). This program also violates contextual noninterference, since client can tell the difference between two secret author lists: say, \([\text{}]\) and \([\text{alice}]\), where client \( \neq \text{client}\).

In Lifty, downgrading does not normally appear in user-written code; rather, the compiler uses it when generating policy checks for self-referential policies (see Sec. 2.4.2 for an example).

Algorithmic type checking. As is customary for expressive type systems, the rules in Fig. 8 are not algorithmic: they require “guessing” appropriate policy predicates for intermediate terms (when applying rules \( T - \text{PRINT} \) and \( T - \text{BIND} \)), as well as the predicate \( r \) in \( T - \lfloor \cdot \rfloor \). Our insight is that we can re-purpose liquid type inference [Cosman and Jhala 2017; Rondon et al. 2008; Vazou et al. 2013], which has been previously used to automatically discover unknown refinements, to also discover these unknown predicates. To this end, our typing rules are carefully designed to respect the restrictions imposed by Liquid Types, such as that all unknown predicates occur positively in specifications. As a result, we obtain fully automatic verification for programs with (decidable) context-dependent policies.

3.4 Contextual Noninterference in \( \lambda^L \)

We want to show that well-typed \( \lambda^L \) programs do not leak information. Traditionally, this property is formalized as noninterference: a leak-free program must take low-equivalent stores to low-equivalent stores. In the presence of context-dependent policies, this traditional definition does not directly apply. Instead, we enforce contextual noninterference, a guarantee similar to that of the Jeeves language [Austin et al. 2013; Yang et al. 2012].

One difference is that our security levels are defined relative to an observer. Hence we want to show that an observer \( o : \text{User} \) cannot tell the difference between two \( o\text{-equivalent} \) stores (those that only differ in locations secret from \( o \)).

More importantly, \( \lambda^L \) policies depend on the values in the store, which brings two major complications. First, a store update can cause a previously secret location to become visible, thereby breaking \( o\text{-equivalence} \). For example, advancing the conference phase to \text{Done} automatically reveals all (previously secret) status locations, which is the intended behavior. Hence, we cannot show that all well-typed \( \lambda^L \) programs preserve \( o\text{-equivalence} \); instead, we show that \text{print}-programs do so, while \text{set}-programs produce stores that agree on all previously visible locations. By combining these two lemmas, we can show that programs in the full language (with multiple \text{print} and \text{set} statements, Appendix C) preserve equivalence on the set of locations that are never revealed.

Second, consider two stores that differ on location \( l \), where \( l \) is \textit{secret in one store and visible in the other}. Should \( o \) be permitted to differentiate between these stores? Following Jeeves, contextual non-interference allows this observation (in other words, these stores are \textit{not} \( o\text{-equivalent} \) by our definition). This decision is motivated by real-world scenarios we are trying to model: for example,
in a social network where my location is only visible to my friends, I don’t care if the attacker can tell that my real location is hidden, as long as they cannot tell whether I’m at home or at work. At a first glance, this might introduce leaks through policy checks: indeed, the attacker now knows that we are not friends. However, such leaks are actually prevented by the $\lambda^T$ type system: if the code passed the type checker, the friendship information must have been visible to the attacker\footnote{This is similar to the LIO [Stefan and Mazières 2014] rule that "the label on the label is the label itself".}.

Definition 3.1 (policy of a type). For a tagged type, $\text{pol}(\langle T \rangle^n) = \pi$; otherwise $\text{pol}(T) = \lambda.\top$.

Definition 3.2 (observable location). For some observer $o$: $\text{User}$ and two stores $\sigma_1, \sigma_2$, a location $l$ is $o$-observable—written $l \in \text{obs}_o(\sigma_1, \sigma_2)$—if it is visible to $o$ at least in one store:

$$\pi[(\sigma_1, o)] \lor \pi[(\sigma_2, o)]$$

where $\pi \equiv \text{pol}(\text{ty}(l))$

Definition 3.3 (observational equivalence). For some observer $o$: $\text{User}$, two stores $\sigma_1, \sigma_2$ are $o$-equivalent—written $\sigma_1 \sim_o \sigma_2$—if they agree on $o$’s output and all $o$-observable locations:

$$\sigma_1[o] = \sigma_2[o] \land \forall l \in \text{obs}_o(\sigma_1, \sigma_2).\sigma_1[l] = \sigma_2[l]$$

Lemma 3.4 (contextual noninterference of expressions). For some observer $o$: $\text{User}$ and two $o$-equivalent stores, $\sigma_1 \sim_o \sigma_2$, an expression $e$ of an $o$-observable type evaluates to the same value: if $e + e :: T$ with $\text{pol}(T)([\sigma_1, o])$ and $\sigma_j, e \rightarrow^* v_j$ for $j \in \{1, 2\}$, then $v_1 = v_2$.

Proof outline. The proof is by induction on the typing derivation (after generalizing to open terms). The only nontrivial case is downgrading: $\Gamma; \sigma \vdash [e] :: \langle \text{Bool} \rangle^n$. We have $\Gamma; \sigma \vdash e :: (\langle \text{Bool} \mid v \Rightarrow r \rangle)^{\lambda.\pi[[s, u]]^{\land r}}$ and $\sigma, e \rightarrow^* v_j$; we need to show $v_1 = v_2$. Suppose this is not the case; then since $\sigma_1 \sim_o \sigma_2$, this means that $e$ reads from unobservable locations, which in turn means that its policy must be violated in both stores. But this can only happen if $\sigma_1[r] = \sigma_2[r] = \perp$ (since $\pi[(\sigma_j, o)]$ holds for $j \in \{1, 2\}$. If that is true, however, then by soundness of refinement types it follows that $v_1 = v_2 = \perp$, a contradiction.

Lemma 3.5 (contextual noninterference for print-statements). For some observer $o$: $\text{User}$ and two $o$-equivalent stores, $\sigma_1 \sim_o \sigma_2$, a print-program $s$ of the form $\text{let } x_1 = e_1 \text{ in } \ldots \text{ print } u y$ produces $o$-equivalent stores: if $\sigma_j, s_j \rightarrow^* s'_j$ for $j \in \{1, 2\}$, then $\sigma'_1 \sim_o \sigma'_2$.

Proof outline. From the type rule for print and soundness of refinement types, we can show that $\pi[(\sigma_j, o)]$ holds for $j \in \{1, 2\}$, and apply Lemma 3.4.

By definition of $o$-equivalence, it follows that $o$ cannot observe the difference between two $o$-equivalent stores through output.

Lemma 3.6 (contextual noninterference for set-statements). For some observer $o$: $\text{User}$ and two $o$-equivalent stores, $\sigma_1 \sim_o \sigma_2$, a set-program $s$ of the form $\text{let } x_1 = e_1 \text{ in } \ldots \text{ set } x y$ produces stores that are equivalent on previously visible locations: if $\sigma_j, s_j \rightarrow^* s'_j$ for $j \in \{1, 2\}$, then $\sigma'_1[o] = \sigma'_2[o] \land \forall l \in \text{obs}_o(\sigma_1, \sigma_2).\sigma'_1[l] = \sigma'_2[l]$.

4 TARGETED SYNTHESIS FOR $\lambda^T$

We now turn to the heart of our system: the type-targeted synthesis algorithm Enforce, shown in Fig. 9. The algorithm takes as input a typing environment $\Gamma$, a program $e$, and a top-level type annotation $T$, and determines whether policy-enforcing code can be injected into $e$ to produce $e'$, such that $\Gamma \vdash e' :: T$. The algorithm proceeds in two steps. First, procedure Localize identifies unsafe terms (line 2), replacing them with type casts to produce a “program with holes” $\hat{e}$ (Sec. 4.1). Second, the algorithm replaces each type cast in $\hat{e}$ with an appropriate patch, generated by the procedure Generate (Sec. 4.2).

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1: \textbf{ENFORCE}(\Gamma, e, T)
2: \hat{e} \gets \textsc{Localize}(\Gamma \vdash e :: T)
3: \textbf{for } (T_e \triangleleft T_a) d \in \hat{e} \textbf{ do}
4: \quad d' \gets \textsc{Generate}([x_0 : T_a], \Gamma, T_e)[x_0 \mapsto d]
5: \quad \hat{e} \gets \hat{e}[\langle T_e \triangleleft T_a \rangle d \mapsto d']

6: \textsc{Generate}(\Gamma_B, \Gamma_G, \langle T \rangle^\pi)
7: \quad \Gamma_B \gets \Gamma_B \cup \mathcal{R}
8: \quad \text{branches} \gets \textsc{SynthesizeAll}(\Gamma_B \vdash ?? :: \langle T \rangle^\lambda^\bot)
9: \quad \text{conds} \gets \textsc{Abduce}(\Gamma_G, ?? \vdash b :: \langle T \rangle^\pi) \text{ for } b \in \text{branches}
10: \quad ((\text{dflt}, c) : \text{guarded}) \gets \text{sort } (\text{branches, conds}) \text{ by } \text{conds} \text{ weakest to strongest}
11: \quad \textbf{if } c = \top \textbf{ then}
12: \quad \quad \text{patch} \gets \text{dflt}
13: \quad \textbf{else fail}
14: \quad \textbf{for } (c, b) \gets \text{guarded} \textbf{ do}
15: \quad \quad \text{guard} \gets \textsc{Synthesize}(\Gamma \vdash ?? :: \langle \{ \nu : \text{Bool} | \nu \leftrightarrow c \} \rangle^\pi)
16: \quad \quad \text{patch} \gets \text{bind}(\text{guard})(\lambda d. \text{if } d \text{ then } b \text{ else } \text{patch})
17: \quad \quad \textbf{return } \text{patch}

Fig. 9. Policy enforcement algorithm

4.1 Fault Localization

\textbf{Type casts.} For the purpose of fault localization, we extend the values of $\lambda^L$ with type casts:

\[ v ::= \cdots | \langle T \triangleleft T' \rangle \]

Statically, our casts are similar to those in prior work [Knowles and Flanagan 2010]; in particular, the cast $\langle T \triangleleft T' \rangle$ has type $T' \rightarrow T$. However, the dynamic semantics of casts in $\lambda^L$ is undefined. The idea is that casts are inserted solely for the purpose of targeting synthesis, and, if synthesis succeeds, are completely eliminated. We restrict the notion of type-safe $\lambda^L$ programs to those that are well-typed and free of type casts.

\textbf{Sound localizations.} The goal of algorithm \textsc{Localize} is to infer a minimal sound localization of the term $e$.

\textit{Definition 4.1 (sound localization).} A term $\hat{e}$ is a sound localization of $e$ at type $T$ in $\Gamma$, if (1) $\hat{e}$ is obtained from $e$ by inserting type casts, \textit{i.e.} replacing one or more subterms $e_i$ in $e$ by $\langle T_i \triangleleft T'_i \rangle e_i$; (2) it is type correct, \textit{i.e.} $\Gamma \vdash \hat{e} :: T$

In particular, note that (2) implies that each $e_i$ has type $T'_i$.

\textit{Lemma 4.2 (Localization).} Replacing each subterm of the form $\langle T_i \triangleleft T'_i \rangle e_i$ in a sound localization of $e$ with a type-safe term of type $T_i$, yields a type-safe program.

This lemma follows directly from (2) and the substitution lemma for refinement types [Knowles and Flanagan 2010]. Crucially, it shows that once a sound localization has been found, patch generation can proceed independently for each type cast.

\textbf{Minimal localizations.} Among sound localizations, not all are equally desirable. Intuitively, we would like to make minimal changes to the behavior of the original program. Formalizing and checking this directly is hard, so we approximate it with the following two properties. A sound
localization is *syntactically minimal* if no type cast can be removed or moved to a subterm without breaking soundness. Picking syntactically minimal localizations leads to patching smaller terms, rather than trying to rewrite the whole program.

Once the unsafe terms are fixed, we can still pick different expected types $T_i$. Intuitively, the more restrictive the $T_i$, the less likely are we to find a patch to replace the cast. A *minimal localization* is syntactically minimal, and all its expected types cannot be made any less restrictive without breaking soundness. In general, there can be multiple minimal localizations, and a general program repair engine would have to explore them all, leading to inefficiency. For the specific problem of policy enforcement, however, there is a reasonable default, which we infer as shown below.

**Inferring the localization.** Algorithm [*Localize*](#) first uses liquid type inference [Cosman and Jhala 2017; Rondon et al. 2008; Vazou et al. 2013] to reduce the problem of checking the source program $e$ against type $T$ to a system of Horn constraints. If the constraints have a solution, it returns $e$ unmodified. Otherwise, [*Localize*](#) first makes sure that all conflicting clauses have been generated by the second premise of $\ll$-Tag2 (implication on policy predicates). If this is not the case, the type error is unrelated to information flow, and is simply reported to the user. If the check succeeds, however, [*Localize*](#) removes those conflicting Horn clauses that were generated by the smallest term, and re-runs the fixpoint solver on the remaining system, inferring strongest solutions for policy predicates.

Next, it resets the non-policy refinements of the removed terms to $\top$ and re-check the validity of the constraints. If the constraints are satisfied, we have obtained a sound and minimal localization (the expected types are the least restrictive because policies are strongest, and other refinements are $\top$). If the constraints are violated, it indicates that the program depends on some functional property of the unsafe term we want to replace. We consider such programs out of scope: if a programmer wants to benefit form automatic policy enforcement, they have to give up the ability to reason about functional properties of sensitive values, since our language reserves the right to substitute them with other values.

### 4.2 Patch Generation

Next, we describe how our algorithm replaces a type-cast $\langle T_e \triangleleft T_a \rangle d$ with a type-safe term $d'$ of the expected type $T_e$, using the patch generation procedure [*Generate*](#) (line 4). At a high level, the goal of this step is to generate a term of a given refinement type $T_e$, which is the problem tackled by type-driven synthesis as implemented in [*Synquid*] [Polikarpova et al. 2016]. Unfortunately, [*Generate*](#) cannot use [*Synquid*] out of the box, because the expected type $T_e$ is not a full functional specification: this type only contains policies but no type refinements, allowing trivial solutions to the synthesis problem, such returning an arbitrary constant with the right type shape.

To avoid such undesired patches, procedure [*Generate*](#) implements a specialized synthesis strategy: first, it generates a list of *branches*, which return the original term redacted to a different extent; then, for each branch, it infers an optimal *guard* (a policy check) that makes the branch satisfy the expected type; finally, it constructs the patch by arranging the properly guarded branches into a (monadic) conditional.

**Synthesis of branches.** In line 8, [*Generate*](#) uses [*Synquid*] to synthesize the set of all terms up to certain size with the right content type, but with no restriction on the policy. Branches are generated in a restricted environment $\Gamma_B$, which contains only the original faulty term and a small set of *redaction functions* $R$. This set is specified by the programmer, and typically includes a “default value” of each type, as well as custom redactions such as $\texttt{mbMap}$ in Sec. 2.4.3. This restriction gives the user control over the space of patches and also makes the synthesis more efficient.
Synthesis of guards. For each of the branches $b$ (which include at least the original term), GENERATE attempts to synthesize the optimal guard that would make $b$ respect the expected type. At a high level, this guard must be logically equivalent to a formula $c$, such that $(1)$ $c \land p \leftrightarrow q$, where $\lambda(s,u).p \doteq \pi$ is the expected policy of the patch, and $\lambda(s,u).q$ is the actual policy of branch $b$; $(2)$ $c$ does not mention the policy parameters $s$ and $u$. This predicate can be inferred using existing techniques, such as logical abduction [Dillig and Dillig 2013]. In particular, GENERATE relies on SYNQUID’s liquid abstraction mechanism [Polikarpova et al. 2016] to infer $c$ for each branch in line 9.

Next, we sort the branches according to their abduced conditions, form weakest to strongest (i.e. in the reverse order of how they are going to appear in the program). In line 11, we check that the first branch can be used as the default branch, i.e. it is correct unconditionally. This property is always satisfied as long as $\Gamma_B$ contains a value $v$ of type $T$.

The main challenge of guard synthesis, however, is that the guard itself must be monadic, since it might need to retrieve and compute over some data from the store. Since the data it retrieves might itself be sensitive, we need to ensure that two conditions are satisfied (1) functional correctness: the guard returns a value equivalent to $c$, and (2) no leaky enforcement: the guard itself respects the expected policy $\pi$ of the patch. To ensure both conditions, we obtain the guard via type-driven synthesis, providing $\langle \{v : \text{Bool} \mid v \leftrightarrow c\} \rangle^\pi$ as the target type.

Lemma 4.3 (Safe patch generation). If GENERATE succeeds, it produces a type-safe term of the expected type $\langle T \rangle^\pi$.

Assuming correctness of SYNTHESIZE and ABDUCE, we can use the typing rules of Sec. 3 to show that the invariant $\Gamma \vdash \text{patch} :: \langle T \rangle^\pi$ is established in line 12 and maintained in line 16. In particular, the type of bound variable $q$ in line 16 is $\{v : \text{Bool} \mid v \leftrightarrow c\}$, hence, then branch is checked under the path condition $c \iff \top$, which implies $\Gamma_G \vdash b :: \langle T \rangle^\pi$.

We would also like to provide a guarantee that a patch produced by GENERATE is minimal, i.e., in each concrete execution, its return value retains the maximum information allowed by $\pi$ from the original term. We can show that, for a fixed set of generated branches, the patch will always pick the most sensitive one that is allowed by $\pi$, since the guards characterize precisely when each branch is safe. Of course, the set of generated branches is restricted to terms of certain size constructed from components in $\Gamma_B$. The original term, however, is always in $\Gamma_B$, hence we are guaranteed to retain the original value whenever allowed by $\pi$.

4.3 Guarantees and Limitations

In this section we summarize the soundness guarantee of type-targeted synthesis in $\lambda^L$ and then discuss the limitations on its completeness and minimality.

Theorem 4.4 (Soundness of type-targeted synthesis). If procedure ENFORCE succeeds, it produces a program that satisfies contextual noninterference.

This is straightforward by combining Lemmas 4.2 and 4.3 with Lemma 3.5.

Completeness. When does procedure ENFORCE fail? As explained in Sec. 4.1, LOCALIZE can fail, when its proposed localization is not safe, which can happen either because the original program contains an error unrelated to information flow, or because the program depends on some functional property of a sensitive value we want to replace. GENERATE can fail in lines 9, 13, and 15. The first failure can happen if the abduction engine is not powerful enough to infer the precise difference between the expected and the actual policy of a branch (this does not happen in our case studies). The second failure indicates that $\Gamma_B$ does not contain any sufficiently public terms; in this case, LIFTY prompts the programmer to add a default value of an appropriate type.
Table 1. Microbenchmarks.

<table>
<thead>
<tr>
<th>Benchmark</th>
<th>Code size (AST nodes)</th>
<th>Time</th>
</tr>
</thead>
<tbody>
<tr>
<td></td>
<td>Original</td>
<td>(LIFTY)</td>
</tr>
<tr>
<td>1 EDAS</td>
<td>79</td>
<td>25</td>
</tr>
<tr>
<td>2 EDAS-Multiple</td>
<td>100</td>
<td>50</td>
</tr>
<tr>
<td>3 EDAS-Self-Ref</td>
<td>100</td>
<td>76</td>
</tr>
<tr>
<td>4 Auction</td>
<td>48</td>
<td>83</td>
</tr>
<tr>
<td>5 Auction-Place-Bid</td>
<td>74</td>
<td>58</td>
</tr>
<tr>
<td>6 Search</td>
<td>82</td>
<td>25</td>
</tr>
<tr>
<td>7 Sort</td>
<td>72</td>
<td>58</td>
</tr>
<tr>
<td>8 Broadcast</td>
<td>25</td>
<td>25</td>
</tr>
<tr>
<td>9 HotCRP</td>
<td>61</td>
<td>29</td>
</tr>
<tr>
<td>10 AirBnB</td>
<td>60</td>
<td>31</td>
</tr>
<tr>
<td>11 Instagram</td>
<td>73</td>
<td>25</td>
</tr>
</tbody>
</table>

Fig. 10. Scalability—N accesses in a single function.

The third failure is the most interesting one: it happens when no guard satisfies both functional and security requirements, indicating that the policy is not enforceable without leaking some other sensitive information. For instance, in the EDAS leak example, if the programmer declared phase to be only visible to the program chair, no program could enforce the policy on status without leaking the information about phase. In this case, LIFTY prompts the programmer to change the policies in a way that respects dependencies between sensitive fields (i.e. to make the policy on status at least as restrictive as the one on phase).

**Minimality.** We would like to show that the changes made by ENFORCE are minimal: in any execution where e did not cause a leak, e’ would output the same values as e. Unfortunately, this is not true, even though we have shown that LOCALIZE produces least restrictive expected types and GENERATE synthesizes minimal patches. The reason is that even the least restrictive expected type might over-approximate the set of output contexts, because of the imprecisions of refinement type inference. In these cases, ENFORCE is conservative: i.e. it hides more information than is strictly necessary. One example is if the state is updated in between the get and the print by calling a function for which no precise refinement type can be inferred. Another example is when the same sensitive value with a viewer-dependent policy is displayed to multiple users. In our case studies, we found that in the restricted class of data-centric applications that LIFTY is intended for, these patterns occur rarely. One approach to overcoming this limitation would be to combine type-targeted synthesis with runtime techniques similar to Jeeves.

5 EVALUATION

To evaluate type-targeted synthesis, we have implemented a prototype \textsc{Lifty} compiler. Our implementation extends the Synquid program synthesizer [Polikarpova et al. 2016]. From Synquid, \textsc{Lifty} inherits ML-style polymorphism, fixpoints, and user-defined data types. As a result, we were able to encode the tagged monad as a library, with $\langle \cdot \rangle$ implemented as a polymorphic data type with a hidden constructor. After inserting the necessary patches, our compiler translates the
resulting LIFTY code into Haskell, and can link it with other Haskell modules. Hence, a possible usage scenario for LIFTY is to serve as a language for the data-centric application core, while low-level libraries can be implemented directly in Haskell.

Using the LIFTY compiler, we implemented (1) a set of microbenchmarks that highlight challenging scenarios and model reported real-world leaks; and (2) three larger case studies based on existing applications from the literature. Our goal is to evaluate the following parameters:

- **Expressiveness of policy language.** We demonstrate that LIFTY is expressive enough to support interesting policies from a range of problem domains, including conference management, course management, health records, and social networks. In particular, we were able to replicate all the desired policies in three case studies from prior work [Swamy et al. 2010; Yang et al. 2016].
- **Scalability.** We show that the LIFTY compiler is reasonably efficient at error localization and synthesis: LIFTY is able to generate all necessary patches for our conference manager case study (450 lines of LIFTY) in about 20 seconds. Furthermore, we show that synthesis times are linear in the number of required patches.
- **Quality of patches.** We compare the code generated by LIFTY to a version with manual policy enforcement; we show that LIFTY allows for policy descriptions to be centralized and concise, and the compiler is able to recover all necessary policy-enforcing code, without reducing functionality.

All programs in our evaluation share a common library, which offers a set of basic functions on primitive types, lists, and the option type, along with a set of monadic combinators for convenient manipulation of tagged values (mapM, filterM, etc). The library also defines a standard set of redaction function, which comprises false, 0, [], "", Nothing, and mbMap. In addition to this set, several programs define custom redaction function, either to designate a default value for a user-defined data type (such as NoDecision in our running example) or to perform nontrivial anonymization (see examples below).

5.1 Microbenchmarks

To exercise the flexibility of our language, we implemented a series of small but challenging microbenchmarks, summarized in Tab. 1. The code of each benchmark, with leak patches inserted by LIFTY, is available in Appendix B.

Benchmarks 1–5 in Tab. 1 correspond to the examples from Sec. 2: EDAS is our running example of the EDAS leak, and the other four are the more challenging scenarios introduced in Sec. 2.4.

Benchmarks 6 and 7 exercise tricky cases of implicit flow through higher-order functions. Search displays the titles of all client’s accepted papers. It is essentially a variant of EDAS, where the implicit flow happens through a filter; LIFTY inserts a patch inside the filter’s predicate.

Sort displays the list of all conference submissions sorted by their score, using a higher-order sort function with a custom comparator. The order of submissions might leak paper scores to a conflicted reviewer. To prevent this leak, LIFTY rewrites the comparator to return a default score is a paper is conflicted with the viewer. Interestingly, this benchmark features a negative self-referential policy for the list of conflicted reviewers: this list is visible only to users who are not on the list. Such policies are not supported by Jeeves, since they are incompatible with its fixpoint interpretation of self-referential policies (Sec. 6); in LIFTY, the semantics of policies is decoupled from their evaluation, hence this example presents no difficulty.
Broadcast sends a decision notification to all authors of a given paper\(^5\). This benchmark tests Lifty’s ability to handle messages sent to multiple users; Lifty infers that all those users are authors of the paper, and hence are allowed to see its status, as long as the phase is Done. An additional challenge is that the list of recipients is *itself sensitive*, since the conference is double-blind; Lifty infers that no additional check is needed, since authors are always allowed to see themselves.

The last three benchmarks model reported real-world leaks. HotCRP models a leak in the HotCRP conference manager, first reported in [Yip et al. 2009], where the conference chair could send password reminder emails to PC members, and then glean their passwords from the email preview. Lifty repairs this leak by masking the password in the preview (but not in the actual email), since the preview is flowing to the chair, while the email is flowing to the owner of the password.

AirBnB models a leak in the AirBnB website, found by one of our students [Voss 2016]. The website redacts phone numbers from user messages (presumably to keep people from going around the site), but phone numbers appear unredacted in message previews. In Lifty we model the AirBnB messaging system by designating the message text visible only to its sender and the site administrator, and introducing a special redaction function `scrubPhoneNumbers`, whose result is additionally visible to the message recipient. With these policies, whenever a message is displayed to the recipient, Lifty inserts a check whether they are the administrator, and otherwise redacts the text with `scrubPhoneNumbers`.

Instagram is inspired by several reported cases, where sensitive social network data was revealed through recommendation algorithms [Hill 2017; Yang 2017]. In particular, if an Instagram account is private, their photos and “following” relations are supposedly only visible to their followers (which have to be approved by the user). Yet, journalist Ashley Feinberg was able to identify the private Instagram account of the former FBI director James Comey, because Instagram mistakenly revealed that James was followed by his son Brien (whose account is public). In Lifty, we model the Instagram “following” relation using a field

\[
isFollowing :: x : \text{User} \rightarrow y : \text{User} \rightarrow \text{Ref} (\text{Bool})^{\lambda}.\text{canSee}(s, u, x) \wedge \text{canSee}(s, u, y)
\]

whose policy requires that both accounts be visible to the viewer (here, `canSee(s, u, x)` holds iff `x` is public, `x = u`, or `u` follows `x`). When the recommendation system attempts to retrieve all accounts followed by Brien Comey, Lifty injects a check that those accounts be visible to the viewer, and otherwise `isFollowing` defaults to false.

### 5.2 Case Studies

We use Lifty to implement three larger case studies: a conference manager and a course manager, both based on the case studies for the policy-agnostic Jacqueline system [Yang et al. 2016], and a health portal based on the HealthWeb example from Fine [Swamy et al. 2010]. The two authors who developed the case studies were not involved in the development of Lifty.

**Conference manager.** We implemented two versions of a basic academic conference manager: one where the programmer enforces the policies by hand and one where Lifty is responsible for injecting leak patches. The manager handles confidentiality policies for papers in different phases of the conference and different paper statuses, based on the role of the viewer. Policies depend on this state, as well as additional properties such as conflicts with a particular paper. The system provides features for displaying the paper title and authors, status, list of conflicts, and conference information conditional on acceptance. Information may be displayed to the user currently logged in.

---

\(^5\)For this purpose we introduce a version of `print` that accepts a tagged list of viewers instead of a single tagged viewer; this operation has to be built in, since Lifty’s statements are unconditional.
record :: rid : RecordId → Ref (RecordId)∧.u=author (s[rid])∨...
    (isPatient s u∧u=patient s[rid]∧¬(withhold u s[rid]))∨...
    (isDoctor s u∧¬(withhold u s[rid]))∨...
    (isPsychiatrist s u∧isTreating u s[rid]∧isPsychiatristRecord s[rid]∧¬(withhold u s[rid]))

isTreating :: x : User → y : User → Ref (Bool)∧.isPsychiatrist s u ∧ u=x

Fig. 11. An excerpt from the policy module of the Health Portal case study.

in or sent via various means to different users. The system contains 888 lines of code in total (524 Lifty + 364 Haskell).

Jacqueline only supports constant default values, but we decided to deviate from the original system to experiment with nontrivial redaction functions. In our version, reviewer names that are hidden for any reason are displayed as ”Reviewer A”, ”Reviewer B”, etc., following common convention. This is implemented by representing each reviewer entry as a pair of (index, name), where the redaction replaces name with ”Reviewer x” according to index.

**Course manager.** We implemented a system for sending grades to students based on their course enrollment and assignment status. An example policy is that a student can see their own scores, whereas instructors can see scores for all of their students.

**Health portal.** Based on the HealthWeb case study from [Swamy et al. 2010], we implemented a system that supports viewing and searching over health records. This case study is interesting because of the complexity of the associated policies; we show some examples from the corresponding policy module in Fig. 11. As you can see, the policy that guards health records is rather complex, which makes the generated enforcement code significantly larger than the size of the original program.

Additional complexity arises because some policies depend on sensitive values. For example, the record’s policy depends on the result of isTreating, but isTreating has a policy of its own, which makes the patients of a psychiatrist visible only to that psychiatrist. Lifty is able to generate and verify a policy check for this case, since in this check isTreating is used only if isPsychiatrist is true.

### 5.3 Performance Statistics

We show compilation times for the microbenchmarks in Tab. 1, and for the case studies in Tab. 2. We break them down into fault localization (including type checking) and patch synthesis. Lifty was able to patch each of the microbenchmarks in under a second, apart from Instagram, which takes around 3.5 seconds. This time is dominated by condition abduction (line 9 of the algorithm Enforce), which scales exponentially with the number of variables in scope and atomic predicates in the relevant policies. For each of the three case studies, Lifty takes under 30 seconds.

**Scalability.** Note that Lifty verifies and patches each top-level function in a program completely independently. Moreover, unlike prior work on program repair, patch synthesis proceeds independently for multiple leaks inside one function. For a stress test, we created a benchmark that performs \(N\) reads (of the same field, for convenience) and then a print to an arbitrary user. Lifty’s job is to patch all of the get locations with a conditional. We show in Fig. 10 that patch generation time is linear in \(N\) (verification is still quadratic in the size of the function, and dominates the compilation time in this case).
In summary, \textsc{Lifty} compilation times are most significantly affected by the complexity of the policies and the number of arguments to the function under repair, while they scale roughly linearly with the overall size of the system. This result suggests feasibility of applying type-targeted synthesis to larger systems.

5.4 Quality of Patches
We compared the two versions of our conference manager (Tab. 2). The column "Original" shows the size of the code, in AST nodes, without any policy enforcement. We also show the cumulative size of policy-enforcing code, both hand-written and generated by \textsc{Lifty}. The size of policy predicates is given as "Policy size". Note that the size of policy-enforcing code often approaches or exceeds the size of the core functionality, which confirms our hypothesis that for data-centric applications, much of the programming burden is in policy enforcement. Manual inspection reveals that while the two versions of policy-enforcing code are syntactically different, they differ in neither functionality nor performance.

6 RELATED WORK
\textsc{Lifty} builds upon several lines of prior work, most notably in sound program synthesis and repair, policy-agnostic programming, static information flow control, and type-directed coercion insertion. Each of these areas has a rich history, but until now they have developed relatively independently.

6.1 Program Synthesis and Repair
\textsc{Lifty} is related to sound program synthesis techniques [Alur et al. 2017; Kneuss et al. 2013; Kuncak et al. 2010; Manna and Waldinger 1980; Polikarpova et al. 2016], which take a formal specification as input and synthesize a provably correct program. These techniques, however, generate programs from scratch, from end-to-end functional specifications, while \textsc{Lifty} injects code into an existing program based on the cross-cutting concern of information flow.

Our problem statement is similar to that of sound program repair [Kneuss et al. 2015], but as we explain in Sec. 2, in the specific setting of policy enforcement, \textsc{Lifty} is able to infer a local specification for each patch, and synthesize all patches independently, which makes it more scalable. There has been prior work on program repair for security concerns [Fredrikson et al. 2012; Ganapathy et al. 2006; Harris et al. 2010; Son et al. 2013], but it does not involve reasoning about expressive information-flow policies, and hence, both the search space for patches and their verification is much less complex.

6.2 Policy-Agnostic Programming
\textsc{Lifty} shares its high-level goal with systems like Hails [Giffin et al. 2012] and Jeeves [Yang et al. 2012], which enforce privacy policies without putting extra burden on the programmer. The main difference is that \textsc{Lifty} enforces policies at compile time, by rewriting the code, while prior work relies on runtime enforcement, which yields nontrivial runtime overheads and makes it difficult to reason about program behavior in the presence of expressive policies.

\textsc{Lifty}’s programming model closely follows the policy-agnostic approach of Jeeves [Austin et al. 2013; Yang et al. 2016, 2012], which factors information flow out from core program functionality, allowing programmers to implement policies as high-level predicates over the program state. Practical policy-agnostic programming relies on two crucial features of the Jeeves semantics. First, Jeeves lets the programmer designate a default value to be used in place of a sensitive value whenever the latter is not visible; this allows the program to do something nontrivial when policy checks fail (instead of just crashing, as in previous runtime enforcement systems). Second, Jeeves supports policies that depend on sensitive values; these policies are ubiquitous in real-world scenarios, and
supporting them directly (rather than having the programmer encode them with labels) is both nontrivial and prevents a whole class of leaks. While powerful, these two features also cause the two main limitations of Jeeves: runtime overhead and unpredictable semantics. With its static approach, Lifty is able to address both of these limitations.

To support default values, Jeeves performs *simultaneous multiple execution* of the program, since it cannot determine a-priori which of the two facets of a sensitive value (the actual value or the default) will be used at the time of output. This execution incurs quite a bit of runtime overhead, as it doubles the number of facets every time the program branches on a sensitive value (for example, sorting a list of sensitive values causes an exponential blowup). Instead, Lifty infers the eventual output context statically, and hence is able to pick the appropriate facet at the time of retrieval. This eliminates runtime overhead, but also makes Lifty conservative in cases where there’s not enough information about the output context available statically, or when a single sensitive value is flowing into different contexts that expect different facets. This imprecision can be addressed by combining static and dynamic techniques, which we leave to future work.

Unpredictable behavior arises in Jeeves because policy predicates are evaluated at run time, and hence, any reference to a sensitive field inside a policy has to be interpreted as a faceted value, in order to prevent leaky enforcement. In the case of self-referential policies this leads to unintuitive fixpoint semantics. For example, if the reviews and the conflicts of a conference submission are both hidden from a conflicted reviewer $u$, then at run time, Jeeves would replace conflicts with an empty list, but reveal the reviews, since $u$ is not a member of conflicts—an empty list, as far as $u$ is concerned. Lifty does not suffer from this limitation: it interprets all policies literally, and prevents leaky checks by verifying the generated patches against the policies. In particular, Lifty would reject a policy as unenforceable if checking this policy causes a leak, while in Jeeves such a policy would silently default to false.

6.3 Static Information Flow Control

The Lifty type system builds upon a long history of work in language-based information flow control [Sabelfeld and Myers 2003]. The most closely related work is Fine [Chen et al. 2010; Swamy et al. 2010] and $F^*$ [Swamy et al. 2011], which also support value-dependent security types and use a monadic encoding of information flow. The key difference is that our system uses (SMT-decidable) predicates as security labels, which supports (1) a direct encoding of Jeeves-style policies, and (2) fully automatic verification and fault localization, crucial for type-targeted synthesis. UrFlow [Chlipala 2010] is the only automated verification system that supports a similar flavor of policies, but it does not provide a sound treatment of self-referential policies. More importantly, none of these approaches address the issue of or programmer burden: they simply prevent unsafe programs from compiling, but do not help programmers write policy-enforcing code.

6.4 Type Coercions and Type Error Localization

Our use of type errors to target program rewriting resembles *type-directed coercion insertion* [Swamy et al. 2009]; in particular, their coercion insertion and coercion generation mechanisms are similar to our fault localization and patch synthesis, respectively, and their coercion set is similar to our set $R$ of redaction functions. The Lifty type system, however, is far more complex than the type systems explored in that work. In particular, the combination of polymorphism and subtyping complicates type error localization (since there are many valid type derivations), while refinements complicate coercion generation (which becomes a refinement type inhabitation problem).

Hybrid type checking [Knowles and Flanagan 2010] can be viewed as coercion insertion for refinement types. In fact, their coercions also amount to wrapping the original value in a conditional, however, in their case both the guard and the alternative branch are straightforward.
Existing work on type error localization for expressive types systems [Loncaric et al. 2016; Seidel et al. 2017; Zhang et al. 2015] is in a more general—yet more forgiving—context of giving feedback to programmers. Our localization technique (removing constraints that make the system unsatisfiable) is similar to [Loncaric et al. 2016], but for our specific purpose we have more information to decide between possible error locations.

7 CONCLUSIONS

This paper makes two main contributions. The first one is type-targeted synthesis: a technique for decomposing the program synthesis problem for a global cross-cutting concern into independent, local synthesis problems. We believe that this is a promising direction to explore in program synthesis research: due to scalability limitations, synthesizing large programs from scratch will likely remain intractable in the near future; hence an important question is how to integrate synthesized code with hand-written code in a meaningful way. Our work answers this question for a particular problem domain; in the future we hope to find other domains amenable to such targeted synthesis.

Our second contribution is a decidable type system for context-dependent policies. With a notable exception of Jeeves, language-based security literature hasn’t explored the implications of security policies that directly depend on the sensitive data they protect: how to verify program against such policies, and what notions of noninterference make sense in this context. In this work, we show that a refinement type system enables fully automatic checking of context-dependent policies, and at the same time, gives them more straightforward semantics than the run-time approach of Jeeves.

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A DESUGARING DO-NOTATION

This is code from Fig. 3 with the do-notation desugared into invocations of bind.

```ocaml
showPaper client p =
  let row =
    bind (get (title p)) (λ t .
      bind (get (status p)) (λ st .
        bind (if st = Accept then get (session p) else return "")
          (λ ses .
            return (t + " " + ses)))))
  print client row
```

B MICROBENCHMARKS

Below you can find the code for all our microbenchmarks with LIFTY-generated patches in gray. The code has been slightly simplified for readability, and the patches have been re-sugared into do notation. We use abbreviations: any ≡ λ (s, u).⊤ and none ≡ λ(s, u).⊥.

**Benchmark 1 (EDAS)**: Show data for paper p to client.

```ocaml
phase :: Ref (Tagged Phase <any>)
title :: PaperId → Ref (Tagged String <any>)
status :: PaperId → Ref (Tagged Status <λ (s,u) . s[phase] = Done>)
session :: PaperId → Ref (Tagged String <any>)
redact {NoDecision}

showPaper client p =
  let row =
    do t ← get (title p)
    st ← do cl ← do x1 ← get phase; x1 = Done
            if cl then get (status p) else NoDecision
    ses ← if st = Accept
            then get (session p)
            else ""
    unwords [t, ses] in
  print client row
```

**Benchmark 2 (EDAS-Multiple)**: Same as (EDAS), but multiple terms need to be patched.

```ocaml
...  -- as in EDAS
authors :: PaperId → Ref (Tagged [User] <λ (s,u) . s[phase] = Done>)

showPaper client p =
  let row =
    do t ← get (title p)
    auts ← do cl ← do x1 ← get phase; x1 = Done
              if cl then get (authors p) else []
    st ← do c2 ← do x2 ← get phase; x2 = Done
              if c2 then get (status p) else NoDecision
    ses ← if st = Accept
```

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11 then get (session p)
12 else ""
13 unwords [t, auts, ses] in
14 print client row

Benchmark 3 (EDAS-Self-Ref): Same as (EDAS-Multiple), but with a self-referential policy on authors.

1 ... -- as in EDAS
2 authors :: p:PaperId → Ref (Tagged [User] \(\lambda (s,u).s[\text{phase}] = \text{Done} \lor u \in s[\text{authors p}]\))
3
4 showPaper client p =
5 let row = do t ← get (title p)
6   auts ← do c1 ← do x1 ← get phase; x1 = Done
7     c2 ← [do x2 ← get (authors p); elem client x2]
8     if c1 ∨ c2 then get (authors p) else []
9   st ← do c3 ← do x3 ← get phase; x3 = Done
10     if c3 then get (status p) else NoDecision
11   ses ← if st = Accept
12     then get (session p)
13     else ""
14     unwords [t, auts, ses] in
15     print client row

Benchmark 4 (Auction) User bid is fully visible to the user themselves and once the phase is Done; partially visible to all users who have bid.

1 phase :: Ref (Tagged Phase \(<\text{any}>\))
2 bid :: x:User → Ref (Tagged (Maybe
3 \(\lambda (s,u).s[\text{bid u}] \neq [] \lor s[\text{phase}] = \text{Done} \lor u = x\))
4
5 showBids client =
6 let showParticipant p =
7   do mB ← do c1 ← p = client
8     c2 ← do x2 ← get phase; x2 = Done
9     if c1 ∨ c2
10     then get (bid p)
11     else do c3 ← [do x3 ← get (bid client); isJust x3]
12         if c3
13         then do x4 ← get (bid p); mbMap (\_ \_ . 0) x4
14         else Nothing
15     maybe "" (liftM show) mB in
16     let out = liftM unlines (mapM showParticipant allParticipants) in
17     print client out
Benchmark 5 (Auction-Place-Bid): Place client’s bid and then show them all bids. Check \( c_3 \) form previous example is omitted, since we know client has bid.

```haskell
... -- as in Auction

placeBid :: World \rightarrow User \rightarrow Int \rightarrow World
placeBid client b =
  let showParticipant p =
      do mb <- do c1 <- p = client
         c2 <- do x2 <- get phase; x2 = Done
         if c1 \lor c2
            then get (bid p)
            else do x4 <- get (bid p); mbMap (\_ \rightarrow 0) x4
      maybe "" (liftM show) mb in
  let out = liftM unlines (mapM showParticipant allParticipants) in
  let newBid = do mb <- get (bid client)
                  setIfNothing b mb in
  set (bid client) newBid
  print client out
```

Benchmark 6 (Search): Show client all their accepted papers. Repairs a leak through a filter.

```haskell
... -- as in EDAS-Self-Ref

showMyAcceptedPapers client =
  let isMyAccepted p =
      do auts <- get (authors p)
         st <- do c1 <- do x1 <- get phase; x1 = Done
             if c1 then get (status p) else NoDecision
         elem client auts \& st = Accept in
  let out =
      do paperIDs <- filterM isMyAccepted allPaperIDs
         titles <- mapM (\p . get (title p)) paperIDs
         unlines titles in
  print client out
```

Benchmark 7 (Sort) Sort papers by their score, which is hidden from conflicted reviewers. Repairs a leak through the order of sorted submission. Contains a negative self-referential policy.

```haskell
title :: PaperId \rightarrow Ref (Tagged String <any>)
conflicts :: p:PaperId \rightarrow Ref (Tagged [User] <\lambda (s,u). \neg (u in s[conflicts p]))>)
status :: PaperId \rightarrow Ref (Tagged Status <\lambda (s,u). \neg (u in s[conflicts p]))>)
sortPapersByScore client =
```
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6 let cmpScore p1 p2 =
7 do x1 ← do c1 ← [do x1 ← get (conflicts p); not (elem client x1)]
8 if c1 then get (score p1) else 0
9 x2 ← do c2 ← [do x2 ← get (conflicts p); not (elem client x2)]
10 if c2 then get (score p2) else 0
11 return (x1 ≤ x2)

Benchmark 8 (Broadcast): Send status notification to authors. Sending message to multiple viewers, viewers are sensitive.
1 authors :: p:PaperId → Ref (Tagged [User] <λ (s,u) . s[phase] = Done \u in s[authors p]>)
2 status :: p:PaperId → Ref (Tagged Status <λ (s,u) . s[phase] = Done ∧ u in s[authors p]>)
3 notifyAuthors p =
4 let st = do cl ← do x1 ← get phase; x1 = Done
5 if c1 then get (status p) else NoDecision in
6 let autos = get (authors p) in
7 printAllAuts (liftM show st)

Benchmark 9 (HotCRP): HotCRP password leak: chair could see other people’s passwords in message preview.
1 userName :: User → Tagged String <any>
2 userPassword :: x: User → Tagged Password <λ (s,u) . u = x>
3 mask :: Tagged Password <none> → Tagged Password <any>
4 redact {mask}
5 sendPasswordReminder u .
6 let preview =
7 do name ← get (userName u)
8 pass ← do cl ← do x1 ← get chair; x1 = u
9 if c1 then get (userPassword u) else mask (get (userPassword u))
10 unwords [name, pass] in
11 let message =
12 do name ← get (userName u)
13 pass ← get (userPassword u)
14 unwords [name, pass] in
15 print (get chair) preview
16 print u message

Benchmark 10 (AirBnB) AirBnB bug: they scrape phone numbers from user messages, but forgot to do so in the preview. This example features a custom redaction function that makes a message

text visible to the recipient, but not completely public. It also showcases expressive functional reasoning with higher-order functions, since correctness depends on the argument of filterM.

```haskell
admin :: Ref (Tagged User <any>)
sender :: m MessageId → Tagged {User | u ≠ σ[recipient m]}<any>
recipient :: m MessageId → Ref {User | u ≠ σ[sender m]} <any>)
text :: m MessageId → Ref (Tagged String <λ(s,u).u = s[admin] ∨ u = s[sender m]>)
scapePhoneNumbers :: m MessageId
  → Tagged String <λ(s,u).u = s[admin] ∨ u = s[sender m]>
  → Tagged String <λ(s,u).u = s[admin] ∨ u = s[sender m] ∨ u = s[recipient m]>
redact {scrapePhoneNumbers}
```

```haskell
viewInbox client =
  let isMyMessage m =
    do to ← get (recipient m)
       to = client in
  let inbox =
    do myMIDs ← filterM isMyMessage allMessageIDs
       messages ← mapM (λm .
         do c1 ← do x1 ← get admin; x1 = u
            if c1 then get (text m) else scrapePhoneNumbers (get (text m))
                ) myMIDs
       unlines messages in
  print w client inbox
```

**Benchmark 11** (Instagram): The James Comey Instagram leak: the follow-relationships of private accounts leak through recommendation algorithms.

```haskell
name :: User → Tagged String <any>
isPrivate :: User → Tagged Bool <any>
isFollowing :: x User → y User → Tagged Bool <λ(s,u).canSee s u ∧ canSee s u y>
inline isPublic s u = ¬s[isPrivate u]
inline canSee s x y = x = y ∨ isPublic s y ∨ s[isFollowing x y]
```

```haskell
follow client newFriend =
  set (isFollowing client newFriend) true
  let shouldFollow u =
    if u = client
      then false
      else do c1 ← do x1 ← get (isPrivate u); ¬x1
              if c1 then get (isFollowing newFriend u) else false in
  let recommendation =
    do uids ← filterM shouldFollow allUsers
       names ← mapM (λu . get (name u)) uids
       unlines names in
  print client recommendation
```
Program Terms

\[\begin{align*}
\nu &::= c \mid \lambda x. e \\
\epsilon &::= \nu \mid x \mid \epsilon \epsilon \mid \text{if } e \text{ then } e \text{ else } e \\
\mid \text{get } x \mid \text{bind } e \epsilon \mid \lfloor e \rfloor \\
\sigma &::= \text{skip} \mid \text{let } x = e \text{ in } s \\
\mid \text{set } x x ; s \mid \text{print } x x ; s
\end{align*}\]

Values

Expressions

Statements

Types

\[\begin{align*}
B &::= \text{Bool} \mid \text{Str} \mid \text{User} \mid \langle T \rangle^\pi \mid \text{Ref } T \\
T &::= \{B \mid r\} \mid T \rightarrow T
\end{align*}\]

Base types

Types

Refinements

\[\begin{align*}
r &::= T \mid \bot \mid \neg r \mid r \oplus \\
\mid r[r] \mid r[r := r] \mid x \mid r r \mid \cdots \\
\text{where } \oplus \in \{=} \mid \wedge \mid \vee \mid \Rightarrow\}
\]

Policy predicates

Constants

\[\begin{align*}
c &::= \text{true} \mid \text{false} \\
\mid (\ldots \text{boolean and string functions}\ldots) \\
\mid l \mid u \mid \text{field}_i
\end{align*}\]

Locations, users

Fields

Fig. 12. Syntax of the core language \(\lambda^L\).

C THE LANGUAGE \(\lambda^L\)

C.1 Syntax of \(\lambda^L\)

The full syntax of \(\lambda^L\) is given in Fig. 12.

C.2 Operational Semantics of \(\lambda^L\)

The full operational semantics of \(\lambda^L\) is given in Fig. 13.

D THE \(\lambda^L\) TYPE SYSTEM

We show the full typing rules in Fig. 14.
Expression Evaluation

\begin{align*}
\frac{e \beta}{e} & \quad \frac{\sigma, (\lambda x: T. e) \ v \ \rightarrow \ [x \mapsto v]e}{e-get} \\
\frac{\sigma, \text{if true then } e_1 \ \text{else } e_2 \ \rightarrow \ e_1}{e-true} & \quad \frac{\sigma, \text{if false then } e_1 \ \text{else } e_2 \ \rightarrow \ e_2}{e-false} \\
\frac{\sigma, \text{bind } v \ (\lambda x: T. e) \ \rightarrow \ [x \mapsto v]e}{e-bind} & \quad \frac{\sigma, [e] \ \rightarrow \ e}{e-downgrade} \\
\frac{e \ \rightarrow \ e'}{e-ctx} & \quad \\
\text{where} & \\
C ::= & \bullet \mid C \ e \mid v \ C \mid \lambda x: T. C \mid \text{get } C \mid \text{if } C \ \text{then } e_1 \ \text{else } e_2 \\
& \mid \text{if } v \ \text{then } C \ \text{else } e_2 \mid \text{if } v \ \text{then } v_1 \ \text{else } \text{Cbind } C \ e \mid \text{bind } v \ C
\end{align*}

Statement Execution

\begin{align*}
\frac{\sigma, \ e \ \rightarrow^* \ v}{\sigma, \ \text{let } x = e \ \text{in } S \ \rightarrow \ \sigma, [x \mapsto v]S} & \\
\frac{\sigma, \text{set } l \ v ; \ S \ \rightarrow \ \sigma[l := v], S}{\sigma, \text{print } u \ v ; \ S \ \rightarrow \ \sigma[u += v], S}
\end{align*}

Fig. 13. $\lambda$ operational semantics.
Enforcing Information Flow Policies with Type-Targeted Program Synthesis

**Well-formedness**

\[
\begin{align*}
\text{WF-} & \quad \Gamma \vdash r : \text{Bool} \\
\text{WF-TAG} & \quad \Gamma \vdash T \\
\Gamma, s : \Sigma, \mu : \text{User} + r & \quad \Gamma \vdash \left\langle T \right\rangle ^{\lambda, r}
\end{align*}
\]

**Subtyping**

\[
\begin{align*}
\text{\langle\text{-Sc}\rangle} & \quad \Gamma \vdash B \triangleleft B' \\
\Gamma \vdash r \Rightarrow r' & \quad \Gamma \vdash T_x \triangleleft T'_x \\
\Gamma \vdash T_x \rightarrow T \triangleleft T'_x \rightarrow T' & \quad \Gamma \vdash T \triangleleft \left\langle T \right\rangle ^{\lambda, r} \\
\Gamma \vdash \left\langle T \right\rangle ^{\lambda, r} \triangleleft \left\langle T' \right\rangle ^{\lambda, r}
\end{align*}
\]

**Expression Typing**

\[
\begin{align*}
\text{\text{T-C}} & \quad \Gamma; \sigma + c :: \text{ty}(c) \\
\text{T-VAR} & \quad x : T \in \Gamma \\
\Gamma; \sigma + x :: T & \quad \Gamma \vdash T_x \rightarrow T \\
\Gamma; \sigma + e :: T & \quad \Gamma \vdash e :: T_
\]

**Statement Typing**

\[
\begin{align*}
\text{T-LET} & \quad \Gamma; \sigma + e :: T \\
\Gamma, \mu : T : T ; \sigma + s & \quad \Gamma; \sigma + \text{let } x = e \text{ in } s \\
\Gamma; \sigma + x :: \text{Ref} & \quad \Gamma; \sigma + \text{print } x_1 x_2 ; s \\
\Gamma, \sigma' :: \left\{ x = \sigma \mid x = x' \right\} ; \sigma' + s & \quad \sigma' \text{ fresh} \\
\Gamma; \sigma + \text{set } x_1 x_2 ; s & \quad \Gamma; \sigma + \text{skip}
\end{align*}
\]

Fig. 14. $\lambda^L$ static semantics.