Abstract

A zero-knowledge proof (ZKP) is a powerful cryptographic primitive used in many decentralized or privacy-focused applications. However, the high overhead of ZKPs can restrict their practical applicability. We design a programming language, Ou, aimed at easing the programmer’s burden when writing efficient ZKPs, and a compiler framework, Lian, that automates the analysis and distribution of statements to a computing cluster. Lian uses programming language semantics, formal methods, and combinatorial optimization to automatically partition an Ou program into efficiently sized chunks for parallel ZK-proving and/or verification. We contribute:

1. A front-end language where users can write proof statements as imperative programs in a familiar syntax;
2. A compiler architecture and implementation that automatically analyzes the program and compiles it into an optimized IR that can be lifted to a variety of ZKP constructions; and
3. A cutting algorithm, based on Pseudo-Boolean optimization and Integer Linear Programming, that reorders instructions and then partitions the program into efficiently sized chunks for parallel evaluation and efficient state reconciliation.

1 Introduction

Zero-knowledge proofs (ZKPs) [24] enable a prover to convince a verifier about the truth of some statement without revealing why. Recent advances by the cryptographic research community have brought a tremendous improvement in the efficiency of general-purpose zero-knowledge (ZK) proving (see [1] for a comprehensive survey), as well as numerous creative applications of ZKPs (e.g., [3, 6, 10, 12, 21]). However, various practical challenges make the adoption of ZK unfeasible for tasks of realistic sizes. A particular problem is the lack of a comprehensive toolkit of programming languages and compilers providing a full suite of effective, intuitive, parallel-aware, and general methods for writing complex statements intended for efficient ZKPs.

To aid practicality, a line of recent work has started to “scale out” ZK protocols: instead of assuming the prover and/or verifier have access to one giant machine, they assume access to a cluster of machines and the capability to distribute the task amongst these
resources. For example DIZK [42] distributes the prover computation of Groth’s ZK protocol [27] by manually partitioning the RICS constraints into equal-sized chunks. zkBridge [43] can distribute the prover computation of Virgo [45] efficiently to multiple servers when the statement already has a high degree of parallelism. Meanwhile, EZEE [44] distributes among multiple machines the prover-verifier computation and communication of an interactive, garbled-circuit-based zero-knowledge proof protocol [30]. Giraffe [38] addresses the challenge of automatically distributing computations in the verifiable outsourcing context by breaking down overly large computations that cannot be outsourced as a whole.

Though efficient in deployment, in order to distribute their chosen ZK protocol over a cluster of servers each of these works require the developer to first both i) modify the ZKP protocols to be amenable to distributed computation; and ii) to manually partition the computation into smaller chunks to be proven in parallel. Such manual processes may be error-prone and lead to suboptimal solutions, especially for statements of real-world complexity.

1.1 Summary of Contributions

In this paper, we design a programming framework – a language Ou and a supporting compiler architecture Lian – enabling programmers to write zero-knowledge statements without thinking about distributed computation. Our compiler automatically and efficiently chunks this program for efficient parallel proving or verification.

(1) Our framework provides a C-like programming language with annotations so that developers, including those without deep knowledge of cryptography, can easily write ZK applications even when using protocols with advanced features.

(2) We develop a compiler and static analysis pipeline, written in OCaml, that calculates the local and communication costs of each program instruction, then automatically finds an efficient way to chunk the corresponding ZK statement. The optimality measure minimizes the sum of the maximum per-machine cost and cumulative communication costs. This analysis can distribute ZK statements incorporating randomized verification.

(3) We prove our system sound, i.e., we show the semantics of the distributed verification of substatements is equivalent to verifying the whole statement centrally.

(4) Empirically, we show that for many statements, including gradient descent and Merkle Tree, our framework can automatically find efficient partitioning of the statements, thus achieving speedup proportional to the number of machines available.

What this paper is not about. Lian does not establish a program’s functional correctness or data-obliviousness, nor the hardness of witness finding for the ZK statement. We assume a program written in our language is sound and secure for its intended ZK application, and is suitably constructed for the target protocol: e.g., we assume the program is data-oblivious if the target is circuit-based.

1.2 Technical Challenges and Solutions

Ou and Lian address a series of challenges arising from each of language design, cryptography, and combinatorial optimization.

---

Figure 1: Running example. We consider this Ou implementation of Freivalds’ algorithm for randomized verification of 100x100 matrix multiplication.

Cryptographic insight. To enable arbitrary partitioning of a statement, the underlying ZK protocol needs to be ‘flexible’: it must allow proving multiple statements of the same witness while ensuring consistency. To this end, our framework focuses on the commit-and-prove paradigm. The prover first commits to their private inputs, and then constructs a proof establishing some (public) relationship among the committed values without revealing them to the verifier. Because the inputs are committed the prover can establish multiple statements on them, all while the verifier is certain their values remain consistent. Any ZKP system can support this capability by proving the consistency of commitments in ZK, but certain protocols have direct support which incurs less cost (e.g., [2, 4, 15, 16, 19, 29, 30, 41]). Given such a ZK backend secure under parallel composition, regardless of its internals, e.g., the use of Fiat-Shamir or random oracle, we can then distribute the proof of any ZK statement in parallel.

During compilation, we first partition the statement into multiple substatements, before the prover precomputes the input/output of each substatement locally. The prover can then in parallel prove to the verifier the correctness of each substatement and then establish the consistency of the input/output between substatements. Although this works in principle, designing a scalable system that automatically distributes the computation with minimum runtime and developer-overhead is still challenging. For example, in ZK data dependency rarely constrains statement partitioning, unlike

```c
# define S 100
# define secparam 128
ploc1 int[S][S] Mmult_plcmx(...) {...}
pvt2 int[S] Mmult_pvtvec(...) {...}
pvt2 int[S] Mmult_pubvec(...) {...}

7 void frvlds(pvt1 int [S][S] M,
     pvt1 int [S][S] M1, pvt1 int [S][S] M2) {
   /* array of verifier controlled randomness */
   pub2 int [S] s = {0};
   int t = 0;
   /* repeat for secparam number of times */
   while (t < secparam) {
     int i = 0;
     while (i < S) {s[i] = v_rand(0, 1); i = i+1;}
     /* compute z = (M1 * (M2 * s)) */
     pvt2 int[S] z = Mmult_pubvec(M2, s);
     pvt2 int[S] z = Mmult_pvtvec(M1, z);
     /* compute q = (M * s) */
     pvt2 int[S] q = Mmult_pubvec(z, s);
     /* check that q and z are the same */
     i = 0;
     while (i < S) {
       assert (q[i] == z[i]); i = i+1;
     }
   }
   return frvlds(commit(M), commit(M1), commit(M2));
}
```
Automatic statement partitioning. In ZK the prover knows both the public and private inputs, and can thus ‘predict’ all intermediate values. As a result, data dependencies can be resolved before proving: computation of \( y = f(g(x_1), x_2) \) can be partitioned into two parallel verification tasks, \( t = g(x_1) \) and \( y = f(t, x_2) \). The prover, with inputs \( x_1 \) and \( x_2 \), just appends \( t \) to form an extended witness. For ordinary computation this parallelization would be infeasible, because \( f \) depends on \( g \). The prover’s ability to ‘locally precompute’ creates new opportunities to maximize parallelism.

To partition effectively, after parsing and typechecking our architecture undertakes two distinct special compilation phases: shallow simulation and deep simulation. The first automatically partitions the ZK statement and then selects variables as witnesses (the appendices to the input, \( i \)) for each of the \( K \) substatements. The parameter \( K \) is user-chosen, and will equal the number of available compute cores in the cluster. Deep simulation then uses the prover’s private inputs to concretely compute the values of those extended witnesses, as well as of any public variables.

Shallow simulation proceeds by using live variable analysis to compute the costs of each instruction in the statement, and then models them as a directed acyclic graph for partitioning by either pseudo-Boolean optimization (PBO) or integer linear programming (ILP). This requires loop unrolling and function inlining as flattening techniques. The resultant ZK substatements are then compiled to the chosen backend protocol implementation for distribution as programs to the participating servers: in our evaluation we use the VOLE-based EMP-toolkit [40]. Lian’s deep simulation then propagates input-dependent values through the statement to concretely compute both public values and the prover’s witnesses for each substatement. To distinguish which values are available during our different simulation phases, we use knowledge levels: \( K_0 \) for (private input-independent) values that are known at compile time, and \( K_1 \) for (private input-dependent) values that are known at distribution time. This way, reusing a proof with different inputs only requires recomputing the deep simulation, and not the shallow simulation or partitioning, where the latter is likely to be the most computationally intensive compilation task.

Supporting randomized verification. As noted, randomized verification can be much more efficient than deterministic verification. However, supporting it during shallow simulation in particular is a challenge. With deterministic verification an oblivious program trace is known at compile time, which allows Lian to analyze the cost of each substatement easily. With randomized verification the program might branch on the at-the-time unknown randomness, preventing the cost from being calculated precisely as whether an expensive statement is entered may depend on a coin flip. Also, privacy imparts an ordering requirement on randomized verification: the verifier’s randomness used in checking a ZK statement can only be revealed after the prover commits to its inputs. This means that the partitioning needs to obey sequential ordering for values that are dependent on dynamic choices of the verifier.

Ou uses a third knowledge level, \( K_2 \), to indicate that the value of a variable can only be known at runtime. The type system enforces non-interference with respect to a security lattice, allowing \( K_0 \)- and \( K_1 \)-leveled values to flow into \( K_2 \)-leveled variables, but not the other way around. The partitioning then treats the \( K_2 \)-leveled variables as one contiguous block in order to estimate worst-case cost and enforce sequential ordering.

Handling large statements. Naively implementing the aforementioned techniques does not scale to large ZK statements. If the cost graph contains too many nodes then finding efficient chunks will be computationally infeasible. Classical graph partitioning and many related variants are NP-hard with efficient (approximate) optimization algorithms known only in certain special cases, and even then often only theoretically [13]. More subtly, a naive representation of each substatement, e.g., as a circuit or R1CS, is necessarily linear in its running time. For large statements, especially those containing unrolled loops or inlined function calls, it is therefore crucial to work with ‘compact representations’ of substatements which are sublinear in the running time.

Ou and Lian support sublinear representations through atomic, a user-provided annotation to indicate that a function should be condensed, rather than flattened, during shallow simulation. By doing so, we contract the component of the cost graph corresponding to the function into a single node of cost cumulative of its constituent instructions. This reduces the size of the graph and makes optimization more practical. Even if the atomic function is called multiple times, only one copy of the declaration exists. This brings the partitioning problem into the realm of practicality for both PBO and ILP solving, despite its inherent hardness.

1.3 Roadmap

Sec. 2 is an overview of the Lian framework’s workflow. Sec. 3 presents the language’s syntax, typing rules, and dynamic semantics. Sec. 4 shows how the shallow simulation unfolds the program into a sequential program, and formally proves the two programs have the same behavior. Sec. 5 then describes how the compiler finds an efficient way to partition the sequential program into multiple chunks and generate a distributed program. Sec. 6 shows how the deep simulation evaluates all dependent data between chunks. It also proves that with these data, the distributed program has the same behavior as if it was run without distribution. Finally, Sec. 7 studies the effectiveness of the framework for distributing the computations of various benchmarks.

2 Lian Design Overview

Figure 2 shows the architecture and workflow of the Lian framework. The user first writes a program in Ou. Upon invocation, the compiler parses and typechecks the program. The Ou type system enforces non-interference over the security lattice given in Figure 3. This lattice specifies non-interference along two dimensions:

- i the public or private nature of the information; and
- ii the knowledge level of the information (i.e., if the information is available at compile time, distribution time or run time)

Accordingly, the lattice defines seven security domains: public \( \text{pub} \times \{ K_0, K_1, K_2 \} \), committed and authenticated \( \text{pvt} \times \{ K_1, K_2 \} \), and prover-local \( \text{plc} \times \{ K_1, K_2 \} \). As syntactic sugar, we write \( \text{pub} \) for \( \text{pub} \times \{ K_0 \} \) and similarly for the rest of the lattice. The variables in the public domain are known by both the prover and the verifier, the variables in the prover-local domain are only known by the
Ou program forms its own independent block in this sequential program: the constituent instructions of functions decorated by the `atomic` keyword are combined into a single block, as are all $K_2$-annotated variables. In both cases, this mapping enforces they be treated as one cumulative ‘instruction’ for the purposes of partitioning. In Figure 1 for example, the suffix of the program beginning with the first `mmult_plcmx` call on L19 will be condensed, thereby guaranteeing the security of the parallel composition in the presence of public randomness.

If the user sets $k = 1$, so that only a single partition is desired, then this sequential program will be directly compiled into the ZK statement as specified by the backend. However, usually $k > 1$ as the developer desires parallelization, and we generally assume this to hold through the remainder of our discussion. The Lian compiler next conducts live variable analysis to track dependencies among the blocks and builds a corresponding directed acyclic graph $G = (V,E)$. An edge $(i, j) \in E$ captures that $j$-th instruction block depends on some information computed in the $i$-th instruction block. For example, in Figure 1 the `commit(M)` on L37 depends on the `mmult_plcmx` block on L36.

In order to obtain efficient partitions, the results of the live variable analysis are used to decorate $G$ with both edge and node labels, each indicating a form of cost to the optimality of any partition. Nodes in $G$ are labeled with their computation cost, which models the time required by the prover and verifier to execute the block’s proving or verification operations in the chosen ZKP protocol backend. Edges are labeled with the cutting cost, which models the communication cost required to reconcile the state of the blocks across compute nodes in the cluster. The latter is linear to the operation’s inbound dependent data. We assume that cutting and computation costs are directly comparable through some scaling factor ($\alpha$), but otherwise omit that detail as it is irrelevant to the function of our compiler.

Lian next uses combinatorial optimization to find the efficient cuts to partition $G$. This optimization problem is most easily understood as an ILP, though as the constituent variables are all Booleans it is amenable to representation as a PBO problem as well.

We formally give the $\text{opt}(\cdot)$ function for this problem, as well as all of its constraints, in Sec. 5.2. Informally, the objective can be understood as the maximum of the cumulative computation costs of each partition, plus the sum of all cutting costs on edges bisected by the $k - 1$ cuts (needed to create the user-specified $k$ chunks). Users can adjust the parameters of the ILP/PBO instances or objective functions to align with different ZK backends and relevant metrics.

---

1The parser accepts unannotated declarations, which are interpreted as `pub0`. 

---

Figure 2: The Lian framework, beginning with an Ou program and ending in binaries for an implementation of a ZKP protocol. Square blocks with light text identify the compiler components: the operations (blue) and their underlying methods (green). Dark text indicates a program or protocol: the Ou program (purple semi-rounded rectangle), compiled substatement binaries (gold rounded rectangles), and the ZKP protocol itself (light blue diamond). Corresponding knowledge levels are displayed as well.

Figure 3: Security Lattice
The partitioned program is compiled into binaries for the chosen ZKP protocol backend. As Lian is designed to use commit-and-prove ZK protocols as a black box without any internal changes of them, our highly adaptable framework can support a wide range of backends, including the NIZK protocols obtained via the Fiat-Shamir transform [5]. These binaries are then sent to any and all individual provers and verifiers. All computation up to this point is based only on public K0 information, and so the binaries may be freely distributed and reused. However, for the programs to run in parallel, the dependent values between partitions must be computed ahead of time. This is the deep simulation. The prover uses their private information to simulate execution of each substatement program, and saves the output values of each one that then become the inputs to others. These output values are the ‘extended witnesses’ of Figure 2. For example, in Figure 1 a cut at L12 would make each entry in M an extended witness. After deep simulation completes, the prover can, with any verifier, execute the ZKP protocol by running their binaries in parallel and connecting each binary to the other party’s running copy over the network (i.e., part0 to part0, part1 to part1, and so on up to partk).

Finally, after all the chunks have been checked in parallel, their inputs and outputs must be reconciled to check consistency of variables passed across the cuts. This step is essential, otherwise the prover could violate soundness by inconsistently computing their extended witnesses, so that, e.g., some variable \( t = 0 \) in one partition but \( t = 1 \) in another. Therefore the runtime of the proving is not only a function of how large each of the partitioned substatements are, but also of how much data must be communicated amongst the compute cores during this consistency checking step – justifying its inclusion in the optimization objective. Lian automatically compiles this reconciliation step into the binsaries themselves, so it requires no additional expertise on behalf of the verifier.

3 Ou Language

3.1 Language Syntax

Figure 4 demonstrates Ou’s syntax. Note that we use \( a^* \) as shorthand for one or more non-terminals separated by commas like \( a_1, \ldots, a_n \), and use \( a^+ \) to represent zero or more non-terminals likewise. A program consists of a collection of struct declaration \( (sd) \), function declaration \( (fd) \), and external function declaration \( (fx) \) statements. There must exist a main function acting as the entry point to the entire program. The syntax is very similar to many imperative languages, except for the addition of security levels \( \ell \) to label primitive types and annotations \( \eta \) to label certain functions.

Each atomic type \( \tau \) is annotated with a security level \( \ell \). Arrays have both a \( \tau \) indicating their elements’ type, as well as an \( \ell \) indicating their access level. The access level can be viewed as the security level of indexes. A struct’s type only mentions its name, and not its fields, so that we can support recursive struct definitions.

There are two different kinds of expressions: L-expressions (\( e \)) and R-expressions (\( \epsilon \)). An L-expression refers to a memory location like a variable \( x \), an array index \( \epsilon[e] \), a struct field \( \epsilon.f \), or a dereferenced pointer \( \epsilon + e \). An R-expression represents a value like an integer, float, or Boolean constant; the calling of a builtin function, array constructor, or struct constructor; or the loading of an L-expression or a pointer to an L-expression. When constructing an

\[
\ell ::= \text{pub0} | \text{pub1} | \text{pub2} | \text{pvt1} | \text{pvt2} | \text{plc1} | \text{plc2} \\
\tau ::= \ell \text{int} | \ell \text{bool} | \ell \text{float} | \tau [\ell \text{num}] | \text{struct} s | \tau + \\
e ::= \ell \text{num} | \ell \text{true} | \ell \text{false} | \text{null} | \text{tt} | (e) \\
\text{[f}'(e) | \ell [\text{e}'] | \{[. \text{label} = \text{e}']\} | e | e \& e \\
e ::= x | e[e] | \ell . e \cdot \epsilon + e \\
c ::= \tau x = e | e = e \\
| \tau x = f(e') \mid c_1 ; c_2 \mid | e | \text{assert} e \\
| \text{if} e \text{then} c_1 \text{else} c_2 | \text{while} e \text{do} c \\
| \text{return} e | \text{return} | \text{break} | \text{continue}
\]

\[
\eta ::= \text{normal} | \text{atomic} | \text{box1} \mid \text{box2} \mid \text{plocal1} \mid \text{plocal2} \\
(\text{sd}) ::= \text{struct} s = \{(\ell \text{label})^{*}\} \\
(\text{fd}) ::= \eta \tau f((\tau x)^{*}) (c) \\
(\text{fx}) ::= \eta \tau f((\tau x)^{*})
\]

\[
\text{prog} ::= (\text{sd})(\text{fd})(\text{fx})
\]

Figure 4: Syntax of the Ou language

array \{ e_0, \ldots, e_{n-1} \}, the user has to specify its access level. The user can also write normal binary and unary operators, including typecasting following the arrows in Figure 3, and all typical arithmetic operations on integers, floats, and Booleans, but Lian will automatically convert them into builtin function calls. For example, given a variable \( x \) of type \text{pvt1 int} and a variable \( y \) of type \text{pub0 int}, \( x + y \) will be automatically rewritten into the expression \text{pvt1 int add}(x, \text{pvt1 int of pub0}(y)).

A command \( c \) can be a variable declaration, an assignment, the calling of a user-defined function, conditional branching, a while loop, a sequence of commands \( c_1 ; c_2 \), a block of commands, returning from a function, or breaking out of a loop.

Note the syntax treats builtin functions and user-defined functions differently. A builtin function \( f \) can be called in expressions freely, while a user-defined function \( f \) has to be received immediately by a variable. This distinction is needed because all builtin functions are pure while user-defined functions can have side-effects. As shallow simulation will reorder computations, having side-effects in expressions can alter behavior. We do still allow calling user-defined functions freely in any expression, but just like in C the programmer cannot assume any execution order as Lian will automatically rewrite the code to extract these calls.

Each function also has an annotation \( \eta \) which enforces restrictions on its body. Most common are normal functions (the parser interprets unannotated functions as normal) which only permit branching on pub0 Boolean expressions. An atomic function is similar, except that shallow simulation will not unfold its body. To branch on a pub1 value, the user must write the code inside a box1 function, while box2 is similar except for additionally allowing branching on pub2 variables. Functions denoted by plocal1 or plocal2 are prover-local functions, which only manipulate plc1 or plc2 values respectively. We call functions denoted by box or plocal sandboxed functions, as the type system enforces restrictions to isolate their effects. We further elaborate in Sec. 3.2.

Before typechecking, Lian first scans the program and collects some top-level definitions: \( \Theta \) maps a struct name \( s \) to its definition \{l_1 : e_1 ; l_2 : e_2 ; \ldots \}. \Lambda^\text{internal} maps a user-defined function name \( f \) to its definition or declaration. When \( f \) is defined, \( \Lambda^\text{internal}(f) \) is like \text{internal} \( \eta (\tau_1 x_1, \ldots, \tau_k x_k) \rightarrow \tau.e \) consisting of its annotation, parameter list, return type, and function body; When \( f \) is
external, \( \Lambda^b(f) \) records only external \( (\eta, (\tau_1, \ldots, \tau_l) \rightarrow \tau) \). Note the language does not support global variables, so the prover can only use external functions to communicate her secret with the program. In addition, the language has builtin functions’ typing information \( \Lambda^b \), which map a builtin function name \( \mathbf{f} \) to its signature \( (\tau_1, \ldots, \tau_n) \rightarrow \tau \). Builtins do not have an explicit function body \( c \), but in Sec. 3.3 we will assign each a math interpretation \( \mathbf{f} \).

In order to ensure the security of ZKP programs, it is crucial for users to label witness variables as \( \text{pvt} \) accurately. Failing to do so can lead to information leakage. Lian offers provably-sound typing rules that enable users to implement secure Ou programs. By implementing a type checker for these rules, Lian can identify and report any errors resulting from incorrect usage of security levels or annotations that violate the typing rules.

### 3.2 Typing Rules

Figure 5 demonstrates the language’s typing rules. Each typing rule relies on the defined strict definitions \( \Theta \) and function definitions \( \Lambda^b \) and \( \Lambda^b \). As these are all fixed we omit them from our written rules for concision. We also rely on a typing environment \( \Gamma \), which is a list of scopes \( \gamma_1, \ldots, \gamma_m \). Each scope \( \gamma_i \) maps variable names to their types. All newly defined variables reside in the first scope \( \gamma_1 \). Note that though the compiler allows the same name to be defined in different scopes, for clarity our discussion assumes that names never conflict.

**Typing orders.** We say \( t_1 \leq t_2 \) if there’s a path from \( t_1 \) to \( t_2 \) in Figure 3. We say a type \( s \) is wellformed if:

- it is an atomic type \( \text{f} \text{t} \text{int} \) or \( \text{t} \text{bool} \),
- it is a struct \( s \) such that \( s \) is defined in \( \Theta \) and the types of all its fields are wellformed;
- it is an array of type \( \tau [\ell n] \) such that \( \tau \) is wellformed and \( \ell \) is no greater than any security level in \( \tau \).

For the last condition the relevant security levels include those of the atomic type as well as the array’s access levels. Wellformedness guarantees the values loaded from arrays are never less secure than the access needed to load them. This restriction is important for both shallow simulation and deep simulation as they need to fully evaluate all \( K_0 \) expressions and \( K_1 \) expressions, respectively. For example, if \( a[e] \) loads a \( \text{pub}0 \) integer from array \( a \) with \( \text{pvt}1 \) access, then shallow simulation does not have enough knowledge to evaluate \( e \), thus it can not know which value is loaded. So although \( a[e] \)’s type is \( \text{pub}0 \), the shallow simulation cannot evaluate it. To make shallow and deep simulation behave as expected, we assume all types in this paper are wellformed.

**Typing expressions.** For L-expressions, \( \Gamma \vdash L e : \tau, \ell \) means in the typing environment \( \Gamma \) an L-expression \( e \) points to a value of type \( \tau \), and the L-expression itself has access level \( \ell \). The access level can help us determine whether an L-expression can be fully evaluated during shallow simulation (\( \text{pub}0 \)) or deep simulation (\( \text{pub}1, \text{pvt}1 \) and \( \text{plc}1 \)) or neither (\( \text{pub}2, \text{pvt}2 \) and \( \text{plc}2 \)). Similarly, for R-expressions, \( \Gamma \vdash R e : \tau \) means an R-expression \( e \) is of type \( \tau \) in the typing environment \( \Gamma \).

Note that in practice the Lian compiler is more forgiving than the language’s strict typing rules, as it performs implicit type conversion along the security lattice’s arrows. For example, an expression \( e \) of type \( \text{pub}0 \text{int} \) can be used as a \( \text{pvt}1 \text{int} \), and Lian will automatically wrap it inside a casting function \( \text{pvt}1\_\text{int}\_\text{of}\_\text{pub}0(e) \).

We also require all pointers to have \( \text{pub}0 \) access level. This restriction is important to the live variable analysis needed for partitioning. Suppose we were to allow a pointer \( p \) to have a \( \text{pvt}1 \) access level: then shallow simulation could not evaluate the pointer itself in order to determine which location is mutated after running \( * p = e \). This would mean all locations in the stack could potentially be the address. Similarly, loading a value from \( * p \) would mean any location could potentially be read. Therefore this restriction on pointer access levels is needed to make the live variable analysis yield meaningful results, rather than just concluding by labeling all variables as live.

**Typing commands.** When typechecking a command \( c \) which constitutes a function’s body, we need its return type \( \tau_r \) to guarantee that each return command in \( c \) yields a valid value. We also need a set of branching security levels \( L_b \) to guard all branching expressions inside \( c \). In normal and atomic functions, all branching expressions have to be \( \text{pub}0 \), but prover-local and boxed functions enforce fewer restrictions.

**Typing a program.** The type checker first checks all struct definitions, and makes sure there are no name collisions between struct names and field names, and then checks each function’s body. For a function definition \( \eta \tau f (\tau_1 x_1, \ldots, \tau_n x_n) (c) \), we apply the command’s typing rule on its body: \( \Gamma, \tau, L_b \vdash [c] : \Gamma \), where \( \Gamma = \{ x_1 : \tau_1, \ldots, x_n : \tau_n \} \) and \( L_b \) is determined by \( \eta \):

\[
L_b = \begin{cases} 
\{\text{pub}0\} & \text{if } \eta \text{ is atomic or normal} \\
\{\text{pub}0, \text{pub}1\} & \text{if } \eta \text{ is box}1 \\
\{\text{pub}0, \text{pub}1, \text{pub}2\} & \text{if } \eta \text{ is box}2 \\
\{\text{plc}1\} & \text{if } \eta \text{ is plc}1\text{al}1 \\
\{\text{plc}1, \text{plc}2\} & \text{if } \eta \text{ is plc}1\text{al}2.
\end{cases}
\]

There are additional restrictions placed on sandboxed functions. We require that \( \text{plc}1\text{al}1 \) and \( \text{box}1 \) functions only modify external data that is \( K_1 \) or \( K_2 \), and only return \( K_1 \) or \( K_2 \) results. Similarly \( \text{plc}2 \) and \( \text{box}2 \) functions can only modify external data that is \( K_2 \), and only produce \( K_2 \) results. The reason is similar to the restrictions on arrays’ access levels: shallow simulation cannot execute \( \text{plc}1\text{al}1 \) or \( \text{box}1 \) functions, so it is infeasible to allow them to modify or return any \( K_0 \) data. If they could, simulation of any other computations that rely on such data could not proceed after the function call.

We also restrict that \( \text{plc}1 \) functions can only manipulate \( \text{plc} \) data. Since all \( \text{plc}1 \) functions only run on the prover’s side they cannot include any \( \text{pub} \) or \( \text{pvt} \) values, as their use would require the verifier’s participation.

To enforce the first restriction, we only need to examine the function’s signature: all pointers passed into \( \text{plc}1\text{al}1 \) and \( \text{box}1 \) functions should only point to \( K_1 \) or \( K_2 \) data, while all pointers passed into \( \text{plc}2 \) and \( \text{box}2 \) functions should only point to \( K_2 \) data. The return types have similar restrictions. Since the language does not support global variables, sandboxed functions can only update external data via the pointers in the arguments. To enforce the second restriction, we need to examine the entire function body to make sure all expressions’ types only have \( \text{plc} \) annotations, and the body only calls other \( \text{plc}1 \) functions.
3.3 Dynamic Semantics

This section defines the dynamic semantics that apply if all secrets are known. The semantics correspond to the program’s behavior in the ideal world where it has enough knowledge to fully evaluate all expressions into values and check if all assertions are satisfied. We will use this semantics as a baseline to prove the simulation and distribution are correct. The semantics rules are not surprising. The detailed definitions are listed in Appendix B.

R-values and L-values. We use \( v \) to denote R-values (or “values” for short) and \( \mu \) to denote L-values. R-values can be atomic values like \( \text{Vnum} \) and \( \text{Vint} \). L-values can be a mapping from indexes to R-values like \( \{0 \mapsto v_0, \ldots, n \mapsto v_n\} \), or a struct \( \text{Vstruct} \) \( m \) where \( m \) is a mapping from label to R-values like \( \{ \text{pub0} \mapsto v_0, \ldots, \text{pubn} \mapsto v_n\} \). An L-value is a reference to a memory location: it can be \( \text{Vvar} \) \( a \) that points to an alias, or \( \text{Vindex} \) \( \mu \) \( l \) that refers to an array \( \mu 's \) index \( i \), or \( \text{Vfield} \) \( \mu \) \( l \) that refers to a struct \( \mu 's \) field \( f \). An alias is used by the stack to disambiguate during recursion, and can be viewed as a memory address.

\[
v := \text{Vnum} | \text{Vint} | \text{Vfloat} | \text{Vbool} | \text{Varray} | \text{Vstruct} | \mu
\]

\[
\mu := \text{Vvar} a | \text{Vindex} m \text{ num} | \text{Vfield} \mu \text{ l}
\]

The program runs on a stack \( \Omega \), i.e., a list of frames \( \omega_1 \cdot \omega_2 \cdots \omega_n \). The last frame \( \omega_n \) stores main’s variables. The stack grows with nested function calls, and the top frame \( \omega_1 \) stores the current function call’s local variables. Each frame \( \omega \) is a map from aliases to values. We use \( \Omega [a \mapsto v] \) to denote adding a mapping \( a \mapsto v \) to the top frame. Whenever a new variable \( x \) is defined, \( \text{Lian} \) allocates a unique alias \( \text{new}(x) \) for it. This is useful when recursion occurs, as the same variable \( x \) might have different instances in the stack, so we cannot bind their values to the same name \( x \). We also use a function \( \text{find}(\Omega, x) \) to find the latest alias allocated for \( x \) in the stack. Note that when \( x \) is defined for the first time, \( \text{new}(x) \) returns \( x \) directly instead of allocating a new alias for it.

Given a stack \( \Omega \), we use a function \( \text{load}(\Omega, \mu, v) \) to load the value referred by \( \mu \) in \( \Omega \), and use a function \( \text{store}(\Omega, \mu, v) \) to compute the updated stack after replacing the value at \( \mu \) with \( v \). They both search the entire stack from top frame to bottom instead of working only with the top frame.

Semantics. The semantics look up the sets of user-defined functions \( A^\Omega \) and built-in functions \( A^\Omega \) to interpret function calls. As they are immutable, for concision we omit them from our rules. Expressions: \([e]_\Omega\) computes E-expression \( e \) in stack \( \Omega \) and yields an L-value, \( [e]_\Omega \) evaluates R-expression \( e \) and yields an R-value. Note \( \text{Lian} \) assigns a mathematical interpretation \( ([]) \) for each builtin function \( f \) which maps a tuple of R-values to an R-value. For example, \( (\text{pvt1_int_add})(\text{Vint} \ pvt1 \ n, \text{Vint} \ pvt1 \ m) = \text{Vint} \ pvt1 \ (n + m) \) is assigned by the compiler for integer addition.

Commands: \( (c, \Omega_1) \rightarrow_C (r, \Omega_2, \beta) \) means that running a command \( c \) from a stack \( \Omega_1 \) yields a sequence of assertion results \( \beta \), a result \( r \) which is either \text{cont} or \text{ret} \( v \), and a changed stack \( \Omega_2 \). Since the command constitutes a function’s body, we use \( \text{ret} \) \( v \) to mean the function returns with \( v \) and \text{cont} to mean it has not yet returned.

4 Shallow Simulation

Shallow simulation executes and unifies the program with \( K_0 \) knowledge to obtain a simplified sequential program. \( \text{Lian} \) will attempt as much simplification as possible, and output all unknown computations as a sequence of commands for later analysis and partitioning. All \text{normal} and \text{atomic} functions will be executed, but only the former’s execution history will be output, while the latter will be output as indivisible calls. As sandboxed functions rely on \( K_1 \) or \( K_2 \) knowledge, they will not be executed in shallow simulation and will be kept in the generated sequential program.

4.1 Shallow Semantics

This section formalizes the shallow simulation’s behavior as a shallow semantics. Figure 6 shows its details. The actual implementation uses several optimizations to speed up the simulation.

Symbolic values. To define the semantics of both shallow simulation and deep simulation, we define symbolic values to represent R-values not known during the simulation. We do so by redefining the R-value constructor from Sec. 3.3 to support Symval variants, and also use symbolic stacks, which store symbolic values.
Builtin functions. During shallow simulation only a subset of the builtin functions can be interpreted as the rest rely on either $K_1$ or $K_2$ knowledge. We use $\langle f \rangle_h$ to denote a shallow simulatable interpretation of a builtin function. It takes in a list of expressions and yields an expression as the result. For example, we have the interpretation $\langle \text{pub0_int_add} \rangle_h(\text{pub0 1, pub0 2}) = \text{pub0 3}$, while $\langle \text{ptv1\_int\_add} \rangle_h(\text{ptv1 1})$ as it relies on $K_1$ knowledge.

Semantics. $[e]_{L_0} = e'$ means simplifying an L-expression $e$ in a symbolic stack $\Omega$ with $K_0$ knowledge results in an L-expression $e'$. These semantics simplify an L-expression instead of evaluating it directly, as the stack may not contain enough information to fully evaluate it to an L-value. Similarly, $[e]_{R_0} = e'$ simplifies an R-expression.

When storing a simplified expression into the stack, we use a partial function $[-]_{L_0}$ to evaluate an L-expression down to an L-value, and a total function $[-]_{R_0}$ to evaluate an R-expression down to a symbolic R-value. $[-]_{R_0}$ is able to be total because it replaces all unknown subexpressions by $\text{SVsym}$. When loading a value from the stack, we use a total function $[-]_{L}$ to lift a symbolic L-value to an L-expression, and a total function $[-]_{R}$ to lift a symbolic R-value to an R-expression and concretize any $\text{SVsym}$ variables using the appropriate L-value. See Figure 19.

We reuse $\text{load}$ and $\text{store}$ to load and store symbolic values to and from the symbolic stack.

$$\text{sload}(\Omega, e) = \begin{cases} \text{load}(\Omega, \mu, \mu'_{L}) & \text{when } [e]_{L_0} = \mu \\ e & \text{otherwise} \end{cases}$$

$$\text{sstore}(\Omega, e) = \begin{cases} \text{store}(\Omega, \mu, [e]_{R_0}) & \text{when } [e]_{L_0} = \mu \\ \Omega & \text{otherwise} \end{cases}$$

When $e$ is not completely known, type wellformedness guarantees that the symbolic value referred to by it must also be unknown, i.e., $\text{SVsym}$ or an array or struct of unknown values. When the so the value to be stored must also be unknown, thus $\text{sstore}$ behaves correctly even when $e$ is not fully evaluated.

Commands: $(c, \Omega_1) \sim_{C} (r, \Omega_2, h)$ means simplifying a command $c$ in a symbolic stack $\Omega_1$ results in: a return status $r$ which is either cont or ret e, a new stack $\Omega_2$, and a sequence of history $h$. We use · to denote an empty history.

Shallow simulation does not have enough knowledge to simulate a sandboxed function, thus it returns a symbolic value $\text{default}(\tau)$ generated from the type $\tau$. Also, the typing rule guarantees even if a pointer is passed into the function, the pointed data must be filled with $\text{SVsym}$. So no matter whether the function has side effect or not, the symbolic stack remains unchanged.

Note that we have added two pseudo-commands $\text{push}$ and $\text{pop}$ into the generated program. They help mark the points where a function is called and returned, and their dynamic semantics are just pushing and popping stacks. Though helpful for the compiler, ignoring them does not change the program’s behavior as the alias mechanism guarantees there will never be name conflicts. We introduce them mainly as machinery to help us prove refinement.

Atomic function calls. The size of a generated program can be very large as it is linear in the execution time. To reduce the challenge for partitioning, we allow users to annotate some functions as atomic in order to contract the function call into one command in the generated program. Atomic annotations are optional for programs and are intended solely for optimization purposes. The effectiveness of an atomic annotation depends on the user’s understanding of the program structure. If an atomic annotation is inappropriate, it may increase the compilation time, but it will not compromise the program’s security.

Their semantics are very similar to normal function calls, but instead of generating a sequence of history, an atomic function call $r x = f(e_1, \ldots, e_n)$ only yields one command in the history: $r a = f(e_1', \ldots, e_n') \ R S_R W S_W$. Here we add a new kind of command to encode atomic function calls, but its dynamic semantics are the same as for normal function calls. The two extra sets $S_R$ and $S_W$ are only used to help perform live variable analysis.

$S_R$ records a set of L-values defined before calling $f$ that may be read during the function call, while $S_W$ records a set of L-values defined before calling $f$ that may be updated during the function call. They are collected during the simulation using functions from Sec. 5.1. $\text{REF}(c)$ computes all L-values that may be read in command $c$; $\text{DEF}(c)$ computes all L-values that may be updated in $c$.

One convenient fact about shallow simulation is that no pointer dereference will appear in a generated program. It is a corollary of a more essential property: shallow simulation fully evaluates all expressions that only need $K_0$ knowledge, including all pointers. In Appendix C, we define normalized expressions to formally describe this property to help prove the correctness of shallow simulation.

4.2 Correctness of Shallow Simulation

We start by defining a refinement function $R$ mapping R-values to symbolic R-values. It replaces all non-\text{pub0} atomic values with $\text{SVsym}$. We overload $R$ to also represent the function that performs the transformations to convert a stack into a symbolic stack.

$$R(\text{Vint } \ell n) = \begin{cases} \text{SVint } \ell n & \text{when } \ell = \text{pub0} \\ \text{SVsym} & \text{otherwise} \end{cases}$$

$$R(\text{Vref } \mu) = \text{SVref } \mu$$

$$R(\text{Varray } \ell a) = \text{SVarray } \ell \cup \{i \mapsto R(a(i))\}$$

$$R(\text{Vstruct } m) = \text{SVstruct } \cup \{i \mapsto R(m(i))\}$$

See Figure 7. We also define a semantics $[r]_{L_2}$ in order to evaluate return statuses.

$$[\text{cont}]_{L_2} = \text{cont}$$

$$[\text{ret } e]_{L_2} = \text{ret } [e]_{L}$$

Theorem 1. If $(c, \Omega_1) \sim_{C} (r, \Omega_2, \beta)$, then there exist $r'$ and $h$ s.t. $(c, R(\Omega_1)) \sim_{C} (r', R(\Omega_2), h)$ and $(h, \Omega_1) \sim_{C} (\text{cont}, \Omega_2, \beta)$ and $[r']_{L_2} = r$.

Figure 7 demonstrates this theorem. Although shallow simulation reordered some computations by simplifying $K_0$ expressions and unfolding the program, this theorem guarantees such reordering does not change the behavior. We prove the theorem in Appendix C.2.

5 Statement Partitioning

The shallow simulation unfolds the original program into a sequential program $h = c_1; \ldots; c_n$. The compiler then tries to find a way to partition these commands into $\kappa$ chunks so that the distributed running time and the communication time are minimized. Sec. 5.1
5.1 Live Variable Analysis

Lian performs live variable analysis on a sequential program $h = c_1; \ldots; c_n$. The goal is to generate a dependency graph $DG = (DV, DE)$ where $DV = \{1 \ldots n\}$ and each edge $(i, j) \in DE$ is decorated with a set of L-values $DEP(i, j)$ that may be read in $c_j$ and may be manipulated in $c_i$. We use ‘may’ rather than ‘will’ because shallow simulation cannot determine all L-expressions’ values. Nonetheless, we can still conduct an over-approximate analysis using some approximation functions. Note $DEP(i, j)$ is only meaningful when $i < j$, so $DEP(i, j) = \emptyset$ when $i \geq j$.

Appendix D defines two utility functions: $REF(c)$ computes the set of L-values that may be read in $c$; while $DEF(c)$ computes the set of L-values that may be manipulated in $c$. Recall that we used these functions in Sec. 4.1.

**Analysis Algorithm.** The program is a sequence of commands $h = c_1; \ldots; c_n$. We start live variable analysis from its end and let $LIVE(n) = \emptyset$. We then scan from the end to the beginning and compute unresolved dependent variables for each command $c_i$.

$$LIVE(i) = \{\mu \mapsto j \mid \mu \notin DEF(c_j) \cup REF(c_j)\} \cup \{\mu \mapsto i \mid \mu \in DEF(c_i)\}$$

$LIVE(i)$ is a mapping from L-values to the commands that may read them in the future. In the first line, if $c_i$ uses an L-value $\mu$
that is later dependent by $c_j$; then we can safely delete it from LIVE($i$) and add that L-value as a dependency between $c_i$ and $c_j$. However, $c_j$ may read other L-values, so in the second line we mark all L-values from REF($c_j$) as unresolved and dependent on $c_j$. Such dependencies will eventually be resolved because every L-value is part of a variable defined in the program, and the command that defines the variable cannot be read. Now we can construct the graph $DG$ by defining the dependency function:

$$\text{DEP}(i,j) = \{ \mu \mid \mu \mapsto j \in \text{LIVE}(i) \land \mu \in \text{DEF}(c_j) \cup \text{REF}(c_j) \}$$

In order to capture all dependencies of and from a command, we denote $\text{DEP}(i,-) := \bigcup_{j=1}^{n} \text{DEP}(i,j)$ and $\text{DEP}(-,j) := \bigcup_{i=1}^{n} \text{DEP}(i,j)$. It follows from our definition that $DG$ is a directed acyclic graph, i.e., $i \geq j \implies \text{DEP}(i,j) = \emptyset$. Any L-value dependent on $c_i$ must either be read or written in $c_i$, any L-value dependent by $c_j$ must be added into the live set during the scan, which can only happen when it is read in $c_j$, so we have the following lemmas:

**Lemma 2.** $\text{DEP}(i,-) \subseteq \text{DEF}(c_j) \cup \text{REF}(c_j)$.

**Lemma 3.** $\text{DEP}(-,j) = \text{REF}(c_j)$.

We need to define a notion of similarity between stacks to reason about the correctness of live variable analysis and program distribution. Given two stacks $\Omega_1$ and $\Omega_2$ and a set of L-values $S$, we say $\Omega_1$ and $\Omega_2$ are similar up to $S$, denoted as $\Omega_1 \approx_S \Omega_2$, if (1) $\forall \mu \in S$, $\text{load}(\Omega_1, \mu) = \text{load}(\Omega_2, \mu)$, and (2) the two stacks have similar structures in terms of $S$, i.e., every alias appearing in $S$ resides in the same frame in both stacks.

Theorem 4 shows the live variable analysis is correct. It says the dependent L-values of $c_j$ are enough to safely execute the command and yield all values that are dependent later. It can be proven using lemma 2 and 3, then prove the values in $\text{REF}(c_j)$ are enough for $c_j$ to execute and produce the desired behavior.

**Theorem 4.** If both $(c_1; \ldots ; c_{j-1}, \Omega) \rightarrow_C (\text{cont}, \Omega_{j-1}, \beta_{j-1})$ and $(c_j, \Omega_{j-1}) \rightarrow_C (\text{cont}, \Omega_j, \beta_j)$, then

$$\forall \Omega'_{j-1}, \Omega'_{j-1} \approx_{\text{DEP}(-,j)} \quad \forall \Omega'_j, (c_j, \Omega'_j) \rightarrow_C (\text{cont}, \Omega'_j, \beta'_j) \land \Omega_j \approx_{\text{DEP}(j,-)} \Omega'_j.$$

### 5.2 Efficient-Cut Search

The final intermediary pass of the compilation partitions the program into $k$ chunks for parallel execution. Recall that after the flattening during shallow simulation (Sec. 4) and the live variable analysis (Sec. 5.1), Lian obtains a sequential program $h = c_1 ; \ldots ; c_n$ and a dependency graph $DG = (DV, DE)$ where each edge $(i, j)$ in $DE$ is decorated with a set of dependent L-values $\text{DEP}(i,j)$. The compiler then extracts a directed acyclic graph $G = (V, E)$ from $DG$ capturing the dependency and the costs. Each node $i$ in $V$ is decorated with $c_i$’s computation cost $PC_i$. An edge $(i, j) \in E$ indicates $c_j$ reads at least one L-value used last by $c_i$, and is decorated with the communication cost $C_{ij}$ and a cuttable bit $B_{ij}$. If there is a $K_0$ L-value in $\text{DEP}(i,j)$, then $c_i$ and $c_j$ must reside in the same chunk, thus $C_{ij} = \infty$ and $B_{ij} = 0$; Otherwise, $C_{ij}$ equals to the accumulated cutting cost of all $K_0$ L-values in $\text{DEP}(i,j)$ and $B_{ij} = 1$.

We use combinatorial optimization constrained by $G$ to find the efficient cutting of $h$ into $k$ chunks ($h_1^{sub}, h_2^{sub}, \ldots, h_k^{sub}$). We naturally write this optimization problem as an integer linear program (ILP). However, the variables (though not the constants) it uses are all Booleans, and therefore the objective can be interpreted as a pseudo-Boolean function and optimized using pseudo-Boolean optimization (PBO) as well.

We begin by defining $ub = \log(\sum_{i=1}^{n} PC_i)$, the length of the cost of the sequential computation. We next define a pair of matrices, $X \in \{0, 1\}^{n \times \alpha}$, indicates the location of cuts: $X_{ij} = 1$ means that a cut is made between the $i$-th and $j$-th instruction block. $X \in \{0, 1\}^{n \times \beta}$ indicates the inclusion of instruction blocks into partitions: $X_{it} = 1$ indicates the $i$-th instruction is in the $t$-th partition for $t \in [k]$.

Finding the cheapest partitioning then reduces to solving the optimization objective of

$$\text{obj}(G) = \min_{X,Y} \left\{ \sum_{(i,j) \in E} Y_{ij} \cdot C_{ij} + \max_{t \in [k]} \sum_{i=1}^{n} X_{it} \cdot PC_i \right\}$$

subject to some consistency constraints on $X$ and $Y$. In order to soundly model this problem for constrained optimization without a min-max objective requires, for PBO in particular, a bit of due care. To convert the original objective into a PBO problem, here we add a bit-vector COP to encode $\sum_{t \in [k]} X_{it} \cdot PC_i$. Our approach is closely related to prior work in the literature on PBO for graph partitioning [17]:

$$\min \sum_{i=1}^{ub} 2^{(i-1)} \cdot \text{COP}_t + \sum_{e \in E} C_e \cdot Y_e$$

s.t.

$$\forall i \in [n] : \quad \sum_{t=1}^{ub} X_{it} = 1$$

$$\forall (i, j) \in E : \quad Y_{ij} \leq B_{ij}$$

$$\forall (i, j) \in E, t \in [k] : \quad X_{it} + X_{jt} \leq 2 - Y_{ij}$$

$$\forall (i, j) \in E, t \in [k] : \quad |X_{it} - X_{jt}| \leq Y_{ij}$$

$$\forall i \in [n], t \in [k] : \quad \sum_{t=1}^{ub} 2^{(t-1)} \cdot \text{COP}_t \geq \sum_{i=1}^{n} X_{it} \cdot PC_i$$

$$\forall i \in [n], t \in [k] : \quad X_{it} \in \{0, 1\}$$

$$\forall t \in [ub] : \quad \text{COP}_t \in \{0, 1\}$$

The first constraint enforces that every instruction appears in exactly one partition, the second that only cuttable edges are cut, and the third and fourth together that if the $i$-th and $j$-th instruction blocks are in different partitions, then any edge $(i, j)$ must be represented in the appropriate cut by $Y_{ij}$. The final four constraints simply enforce the Boolean character of the variables.

The fifth constraint is the most nuanced. It enforces that

$$\text{COP}_t = \text{the } t\text{-th bit of } \max_{t \in [k]} \sum_{i=1}^{n} X_{it} \cdot PC_i$$

so that the computation in the objective function is correct. This relationship can be encoded much more simply in a general ILP, but is necessary for PBO compatibility. Notice that no constraints force the partitions to be contiguous: even though $h$ is a sequential program, the $h_i^{sub}$ can have a much more complex structure, at the
cost of additional extended witnesses and so more communication during the consistency checking.

With this encoding, Lian hands this optimization problem to either a PBO or ILP solver (we use the Gurobi PBO solver in our benchmarks). The solver’s output may not always be optimal in practice, as it is limited by the available computational resources. Consequently, the corresponding cut scheme we employ may also be suboptimal. The compiler then converts the result into a function chunk(−) such that chunk(i) = t if Xi(t) = 1 in the solver’s result.

FACT 1. chunk(i) ≠ chunk(j) ⇒ no K2 L-value in DEP(i, j).

This follows directly from the PBO constraints.

5.3 Distributed Program

Given a sequential program h = c1; . . . ; cn from the shallow simulation, a dependency graph DEP(−, −) from live variable analysis, and a partition chunk(−) from efficient-cut search, the compiler generates k programs for distributed execution among k pairs of prover and verifier computation cores.

The t-th chunk is a sequential program sync(c1; . . . ; cn; . . . ) such that t = chunk(t1) = chunk(t2) = . . . and t1 < t2 < . . . .

We denote a communication between the prover and the verifier by syncφ. These communication rounds are used to share K0 or authenticate K1 data that are needed by cij, i.e., data pointed by L-values in χchunk(i)∪DEP(i, ti), as necessary. Deep simulation (Sec. 6) computes all this dependent data and distribute it to all of the prover cores. So each pair of prover core and verifier core can work together to do a consistency check. The goal is to make sure all data provided by the prover cores in syncφ are indeed the same as produced by their dependent commands in other chunks. Note the deep simulation only computes K0 and K1 values, but Fact 1 guarantees all data needed for multipole chunks does not contain K2 information, thus the deep simulation is enough.

6 Deep Simulation

The deep simulation executes the program with K0 and K1 knowledge and evaluates all K0 and K1 values. The goal is to compute all data that are dependent between different chunks in the distributed program. It is similar to shallow simulation, except that (i) the definitions of all external functions are known, (ii) the knowledge level is K1 instead of K0, and (iii) the purpose is to evaluate all K0 and all K1 values instead of unfolding the program.

6.1 Deep Semantics

This section formalizes deep simulation as a deep semantics. We use the same symbolic values and stacks as in Sec. 4.1. We still use −[−],[−]R to lift R-values and L-values to R-expressions and L-expressions. But when converting in the reversed direction, we use −[−]L and −[−]R, instead of −[−]L and −[−]R. The only difference is when converting atomic expressions. For example, [t n]L = SVInt t n when t ≤ pInt. We also define dload and dstore based on these operations just like sload and sstore.

The semantics for L-expressions [[e]]L and R-expressions [[e]]R are also similar except for using −[−]L and −[−]R to interact with the symbolic stack, allowing dereference expression to be not fully evaluated, and use [[e]]1 to interpret builtin functions. [[ ]1 is a superset of [[ ]0 by adding more functions that can be computed with K1 knowledge. For example, (pvt1_op_t_add)1(pvt1 n, pvt1 m) = pvt1 (n + m). Appendix E has their detailed definitions.

Commands’ semantics is also similar. (c, k1) −→C (r, k2) means running command c from stack k1 results in a return status r which is either cont or ret e and a new stack k2. Note it no longer generates a history as the purpose of the deep simulation is merely computing K0 and K1 values. Also, it executes normal, atomic, plocal1 and box1 functions while ignores plocal2 and box2. The detailed rules are in appendix E.

6.2 Correctness of Deep Simulation

Similar to shallow simulation, we define a refinement function Q that converts R-values to symbolic R-values. This time we only turn all values with pInt2, pvt2 and pIntc security levels into SVsym. This function is also overridden to apply on stacks.

\[
\begin{align*}
Q(\Omega_1) & \xrightarrow{c} Q(\Omega_2) \\
\beta, r & \mapsto \Omega_2 \\
\Omega_1 & \xrightarrow{c} r'
\end{align*}
\]

Figure 8: Deep Semantics Refinement

THEOREM 5. If (c, k1) −→C (r, k2, β) then there exists r' where (c, Q(Ω1)) −→C (r', Q(Ω2)) and [r' ]R = r.

The proof is very similar to theorem 1 as their semantics are similar except for the deep semantics is a bit simpler.

6.3 Correctness of Distribution

Sec. 5.3 shows the distributed program, and Sec. 6.1 evaluates all dependent data that are needed in the sync steps. We now prove these data are sufficient for the distributed program to terminate with the same behavior as if running without distribution.

Given a sequential program c1; . . . ; cn generated by the shallow simulation and an initial stack Ω0, suppose the program runs like (c1, Ω1) −→C (cont, Ω2, βj) without cutting. We focus on the t-th chunk and align its execution like in Figure 9. Each command cj corresponds to two stacks Ωj pre and Ωj post. If chunk(j) = t, we first run syncφ to add external dependent values into Ωj pre, then run cj; Otherwise, the stack remains unchanged.

\[
\begin{align*}
\Omega_{j-1} & \xrightarrow{c_j} \Omega_{j} \\
\Omega_{j-1} & \xrightarrow{\text{cont, } \beta} \Omega_{j+1} \\
\Omega_{j-1} & \xrightarrow{c_j \text{ cont, } \beta} \Omega_{j+1} \\
\Omega_{j-1} & \xrightarrow{c_{j+1} \text{ cont, } \beta} \Omega_{j+1} \\
\end{align*}
\]

Figure 9: Execution in a Chunk

We need to prove that the following theorem holds, then theorem 4 ensures that cj terminates with the same behavior.

THEOREM 6. ∀j ∈ chunk−1(t), Ωj pre ≅ S Ωj−1 where S = DEP(−, j).
In appendix F, we prove this theorem by first defining a set of L-values \( \phi_j \) that are internally dependent in the \( t \)-th chunk after \( c_j \), then prove the stack in the uncut execution is always similar to the \( t \)-th chunk’s local stack up to this set. So the internal dependency is handled, we then use theorem 5 to handle the external dependency happening in \( \text{sync} \), combining them together proves the theorem.

7 Evaluation

7.1 Evaluation Setup and Metrics

We use three metrics to evaluate the effectiveness of our tool:

- **Compilation Time** is the time used for shallow simulation, data dependency analysis, and efficient-cut search. PBO solving during the compilation is done via the Gurobi [28] PBO solver. We do not include the time cost of deep simulation in the compilation time, but deep simulation takes less than 0.5% of the compilation time for all examples.

- **Effective Ratio** of a distributed zero knowledge program is defined as the quotient of the cost of the sequential execution (total costs of all instructions) by the cost of the distributed execution (as the objective function in Sec. 5). When a program is distributed to \( \kappa \) chunks, the closer the effective ratio is to \( \kappa \), the better the distribution scheme is. We note that the optimal effective ratio may be smaller than \( \kappa \), depending on the statement.

- **Execution Time** refers to the end-to-end performance. We compile Ou code to a VOLE-based ZK implementation (EMP [40]) and measure the execution time of the resulting distributed programs. We opt to use VOLE-based ZK backend as one of our motivations is to work with large-scale applications. This choice allows us to handle large statements without requiring costly machines with large memory. In addition, our implementation can be easily extended to support a rich set of ZK protocols, since EMP can be connected to SIEVE-IR [37].

To evaluate our tool, we focus on two programs with different features. All implementations in our language Ou can be found in the supplementary material. As the effectiveness of Lian also depends on the structure of the input program, we further provide visualization for the program structure of gradient descent and merkle tree in the appendix G.

(1) **Gradient descent (GD)** is an optimization algorithm for finding the optimal model in machine learning. It works by iteratively updating the parameters of a model in the direction of the negative gradient of the cost function concerning those parameters. We implement gradient decent for logistic regression with 10 features.

(2) **Merkle tree (MT)** is a type of data structure used to verify the integrity of large amounts of data. It breaks the data into smaller blocks, then recursively hashes each pair of blocks to create a new set of parent nodes using a cryptographic hash function. By proving the correct computation of the Merkle tree, the prover can prove they have the original data without revealing the actual data to the verifier [33].

7.2 Evaluation of the Partition Quality

First, we demonstrate the changes in the effective ratio of the distribution scheme found by Lian for GD and MT program changes with the size of programs and the number of cuts. GD takes a set \( \{(x, y)\} \) as input, where \( x \) is a vector of size \( M \). Throughout this section, we fix \(|\{(x, y)\}|\) to be 20 and \(|x|\) to be 10.

The effective ratio of Lian is affected by the length of the program being distributed. The length of GD and MT programs are decided by the number of iterations and the number of input blocks, respectively. As shown in Figure 10b and Figure 11b, Lian can find the distribution scheme that achieves the almost optimal effective ratio for large programs for both GD and MT cases.

Figure 10a and Figure 11a demonstrate that the effective ratio will increase with the number of machines rising when the number of partitions is relatively small. Moreover, the results present an upper bound of the effective ratio of the distribution scheme for a given program. For example, the effective ratio for the GD programs involving 80 iterations reaches 40 and stays unchanged regardless of the increasing number of machines. Nevertheless, the results show that our tool can utilize a high degree of parallelization.

7.3 Effectiveness of Atomic Annotation

In this section, we study the impact of annotation to the compilation time and effective ratio. For all examples, we set the time limit of the PBO solver to 10 minutes. In the program for GD, there is a function called `update_all()`. This function updates the model’s parameter using all data points and is called for each iteration. There is a function called `sha256_node()` that computes the hash of internal blocks in the MT program. We compare the compilation time and effective ratio of GD programs (or MT programs) with and without `update_all()` (or `sha256_node()`) being atomic while changing \( \kappa \) and the number of iterations.
The results of our evaluation show that the effects of adding atomic annotations depend on the structure of the programs. Figure 12 and Figure 13 show a significant rise in the compilation efficiency when adding annotation for both GD and MT programs. The compilation time decreases by 1X to 60X with atomic annotations. The effective ratio estimation also decreases, especially when the constrained time is far less than what is needed to search for an efficient partition. Figure 13 demonstrates an example of this case. We measure the effective ratio while setting the time limit of the PBO solver to 600 seconds. The effective ratio for the annotated program is smaller than the unannotated one for GD programs: adding atomic annotation rules out a part of feasible solutions which may contain the optimum.

Sometimes, adding annotation could even improve the effective ratio, especially when the constrained time is far less than what is needed to search for an efficient partition. Figure 13 demonstrates an example of this case. We measure the effective ratio while setting the time limit of the PBO solver to 600 seconds. The effective ratio of the MT program for 32-block input with annotation is even better than the one without annotation when the number of cuts increases. One reason behind this result is the MT program structure. The solution found for the annotated program is also optimum for the unannotated one. Furthermore, with a part of the functions being atomic, the search space of the PBO solving essentially shrinks.

For readers who are interested in additional information, we have included the results of MT programs for various input sizes in the appendix.

### 7.4 End-to-End Running Time

Our implementation is end-to-end: it can compile an Ou program to k pieces of EM circuits. Most of the compiler’s implementation is not tied to EMP, so it is flexible to connect with other backends.

We test the running time of the distributed circuits generated from GD and MT. The testing machine is an AWS m5.large instance with 2vCPU and 8GB memory. For GD, we fix the number of dataset to be 10, the size of each dataset to be 100 and iterate 200 rounds.

For MT, we fix the data length to be 256. Table 1 shows with different number of machines, the slowest part’s running time and the overall speedup comparing to uncut execution time. We also list the effective ratio estimated by the solver. For GD, the effective ratio approximates the real speedup very well; For MT, the effective ratio becomes inaccurate as each chunk’s runtime reduces. We think the reason is mainly because the underlying circuits need a constant setup time (around 4 seconds). After subtracting it, the speedup becomes close to the effective ratio.

### Acknowledgement

This work is supported by NSF awards CCF-2106845, CCF-2131476, CCF-1763399, CNS-2016240, CNS-2236819, CCF-2219995, DARPA under Contract No. HR001120C0087, and the Office of Naval Research (ONR) of the United States Department of Defense through an National Defense Science and Engineering Graduate (NDSEG) Fellowship. The views, opinions, and/or findings expressed are those of the author(s) and should not be interpreted as representing the official views or policies of the Department of Defense or the U.S. Government.

### References


### Table 1: End-to-end running time of our framework.

<table>
<thead>
<tr>
<th></th>
<th>runtime GD</th>
<th>runtime speedup GD</th>
<th>effective ratio GD</th>
<th>runtime MT</th>
<th>runtime speedup MT</th>
<th>effective ratio MT</th>
</tr>
</thead>
<tbody>
<tr>
<td># Machines</td>
<td></td>
<td></td>
<td></td>
<td></td>
<td></td>
<td></td>
</tr>
<tr>
<td></td>
<td>1</td>
<td>5</td>
<td>10</td>
<td>20</td>
<td>40</td>
<td></td>
</tr>
<tr>
<td># Effective Ratio</td>
<td></td>
<td></td>
<td></td>
<td></td>
<td></td>
<td></td>
</tr>
<tr>
<td></td>
<td>1.00</td>
<td>4.69</td>
<td>9.57</td>
<td>17.90</td>
<td>31.12</td>
<td></td>
</tr>
<tr>
<td></td>
<td>1.00</td>
<td>4.99</td>
<td>9.95</td>
<td>19.99</td>
<td>39.8</td>
<td></td>
</tr>
<tr>
<td></td>
<td>1.00</td>
<td>3.63</td>
<td>5.27</td>
<td>6.19</td>
<td>7.66</td>
<td></td>
</tr>
<tr>
<td></td>
<td>1.00</td>
<td>4.95</td>
<td>9.75</td>
<td>19.47</td>
<td>38.63</td>
<td></td>
</tr>
</tbody>
</table>

Figure 12: Compiling takes significantly greater time for GD programs when the function update_all(\(\phi\)) is not annotated with the atomic keyword. On the other hand, atomic annotation rules out some feasible solutions, making the distribution scheme slightly suboptimal.

Figure 13: Adding annotation improves both the effectiveness of both compilation and the effective ratio of the MT program for 32-block input.
A Example

Figure 14 shows a code snippet of the Merkle tree we used in the benchmark. The prover uses a binary Merkle tree to compute the hash of $N$ private data blocks. The merkle function computes the hash of a segment of data blocks. Then the verifier calls verify to verify the hash. When the tree is a leaf node, merkle calls sha256_leaf to fetch a data block and hash it; When the tree is a non-leaf node, merkle recursively computes its two subtrees hashes, then use sha256_node to compute the merged hash. The input has 4 data blocks, and we cut the code into 2 fragments.

Here we annotate sha256_node and sha256_leaf as atomic, and load_block as plocal1. So during the shallow simulation, sha256_node and sha256_leaf will not be unfolded, and load_block will be ignored and executed during the deep simulation.

Shallow simulation executes the program with $K_0$ knowledge, yielding a sequence as Figure 15. The recursive function merkle is unfolded. Each variable is annotated with a unique number to distinguish it from others with the same name. Line 1-5 computes the hash of data block 0 and 1, line 6-11 computes the hash of data block 2 and 3, line 13-14 merges them, then line 15 verifies it. Notice in Figure 14 merkle, the recursion condition is a pub8 expression. This is necessary for the unfolding to proceed during the shallow simulation phase. If there is any non-pub8 expression involved in deciding the control flow, then 1) the shallow simulation does not have enough knowledge to evaluate that expression so it can not decide which branch to take, and 2) even if it uses the prover’s knowledge to evaluate that expression, the information will be leaked to the verifier via the generated programs.

Then live variable analysis generates a dependency graph as in Figure 16. The ILP/PBO solver decides to cut it between lines 5 and 6, so the only dependency is hash5. Now the compiler can already generate two programs as in Figure 17. We only show the produce and sync functions here. Note produce_5 in chunk 1 records a digest of all variables that are produced by Figure 15’s line 5 for consistency check in the end. Here it is hash5. Similarly, sync_5 records a digest of all variables dependent by Figure 15’s line 6 after assigning values to them. But we do not know hash5’s value yet, so we mark it using a question mark. The code can already be distributed to individual verifiers, who will count on the provers to supply encrypted hash5 during the runtime. But the provers need to know these values before interacting with the verifiers. So they perform deep simulation on Figure 15’s program based on the secret input and find out hash5’s value when reaching line 5.
/* compute merkle hash of block[left, ..., right-1] */
3 pvt1 int[8] merkle(int left, int right) {
4   if (left + 1 >= right) {
5     pvt1 int[8] hash = sha256_leaf(left);
6     return hash;
7   }
8   else {
9     int mid = (left + right) / 2;
10    pvt1 int[8] hashL = merkle(left, mid);
11    pvt1 int[8] hashR = merkle(mid, right);
12    pvt1 int[8] hash = sha256_node(&hashL, &hashR);
13    return hash;
14  }
15 }
16
17 plocal1 plc1 int[64] load_block(int block_id) { ... }
18
19 /* compute the hash of a block */
20 pvt1 int[8] sha256_block(pvt1 int[64] &data) { ... }
21
22 atomic void verify(pvt1 int[8] * hash) { ... }
23
24 /* compute the hash of a data block */
25 atomic pvt1 int[8] sha256_leaf(int block_id) {
26   /* prove loads and commits a block */
27   plc1 int[64] plc_block = load_block(block_id);
28   pvt1 int[64] block = {0};
29   for int i = 0; i < 64; i = i + 1;
30     block[i] = commit(plc_block[i]);
31   }
32   pvt1 int[8] hash = sha256_block(block);
33   return hash;
34 }
35
36 /* compute the hash of two merged hashes */
37 atomic pvt1 int[8] sha256_node(int left, int right) {
38   pvt1 int[8] *hashL, pvt1 int[8] *hashR) {
39   /* merge two hashes */
40   for int i = 0; i < 8; i = i + 1;
41     block[i] = hashL->[i];
42     block[i+8] = hashR->[i];
43   }
44   for int i = 16; i < 64; i = i + 1;
45     block[i] = 0;
46   } /* compute the hash of the merged results */
47   pvt1 int[8] hash = sha256_block(block);
48   return hash;
49 }
50
51 /* chunk 1 */
52 pvt1 int[8] hash1 = sha256_leaf(0);
53 pvt1 int[8] hashL1 = hash1;
54 pvt1 int[8] hashR0 = hash3;
55 pvt1 int[8] hash5 = sha256_node(&hashL1, &hashR0);
56 pvt1 int[8] hashL2 = hash3;
57 pvt1 int[8] hash7 = sha256_leaf(2);
58 pvt1 int[8] hash3 = sha256_leaf(1);
59 pvt1 int[8] hashR2 = hash11;
60 pvt1 int[8] hashR3 = sha256_node(&hashL3, &hashR3);
61 pvt1 int[8] hash11 = sha256_node(&hashL3, &hashR3);
62 pvt1 int[8] hashR2 = hash11;
63 pvt1 int[8] hashL3 = sha256_node(&hashL0, &hashR2);
64 pvt1 int[8] hash0 = hash13;
65 verify(&hash0);
66
/* chunk 2 */
67 pvt1 int[8] hash5;  
68 plocal1 plc1 int[64] load_block(int block_id) { ... }
69
70 void sync_6() {
71   add_in_digest(1, hash5);
72   add_out_digest(2, hash5);
73   } /* sync_6 */
74
75 /* compute merkle hash of block */
76 pvt1 int[8] merkle(int left, int right) {
77   if (left + 1 >= right) {
78     pvt1 int[8] hash = sha256_leaf(left);
79     return hash;
80   }
81   else {
82     int mid = (left + right) / 2;
83     pvt1 int[8] hashL = merkle(left, mid);
84     pvt1 int[8] hashR = merkle(mid, right);
85     pvt1 int[8] hash = sha256_node(&hashL, &hashR);
86     return hash;
87   }
88 }
89
90 /* compute the hash of two merged hashes */
91 atomic pvt1 int[8] sha256_node(int left, int right) {
92   pvt1 int[8] *hashL, pvt1 int[8] *hashR) {
93   /* merge two hashes */
94   for int i = 0; i < 8; i = i + 1;
95     block[i] = hashL->[i];
96     block[i+8] = hashR->[i];
97   }
98   for int i = 16; i < 64; i = i + 1;
99     block[i] = 0;
100    } /* compute the hash of the merged results */
101   pvt1 int[8] hash = sha256_block(block);
102   return hash;
103 }
104
105 unit main() {
106   pvt1 int[8] hash = merkle(0, N);
107   verify(&hash);
108   return;
109 }
110
111 /* sync_6 */
112 hash5 = 7;
113 add_in_digest(1, hash5);
114
115 /* compute the hash of a block */
116 pvt1 int[8] sha256_leaf(int block_id) {
117   /* prove loads and commits a block */
118   plc1 int[64] plc_block = load_block(block_id);
119   pvt1 int[64] block = {0};
120   for int i = 0; i < 64; i = i + 1;
121     block[i] = commit(plc_block[i]);
122   }
123   pvt1 int[8] hash = sha256_block(block);
124   return hash;
125 }
126
127 unit main() {
128   pvt1 int[8] hash = merkle(0, N);
129   verify(&hash);
130   return;
131 }
132
133 #define N 4
B Detailed Semantics

FACT 2.

\[(c_1; c_2); c_3, \Omega) \triangleright_C (r, \Omega', \beta) \iff (c_1; (c_2; c_3), \Omega) \triangleright_C (r, \Omega', \beta)\]

Proof.

(1) Suppose that \((c_1, \Omega) \triangleright_C \text{ (ret } v, \Omega_1, \beta)\). Then \((c_1; c_2, \Omega) \triangleright_C \text{ (ret } v, \Omega_1, \beta)\), so \(((c_1; c_2); c_3, \Omega) \triangleright_C \text{ (ret } v, \Omega_1, \beta)\). If follows as well that \((c_1; (c_2; c_3), \Omega) \triangleright_C \text{ (ret } v, \Omega_1, \beta)\).

(2) Suppose both that \((c_1, \Omega) \triangleright_C \text{ (cont, } \Omega_1, \beta_1)\) and \((c_2, \Omega_1) \triangleright_C \text{ (ret } v, \Omega_2, \beta_2)\). Then \((c_1; c_2, \Omega) \triangleright_C \text{ (ret } v, \Omega_2, \beta_1, \beta_2)\), so \(((c_1; c_2); c_3, \Omega) \triangleright_C \text{ (ret } v, \Omega_2, \beta_1, \beta_2)\). Also \((c_2; c_3, \Omega_1) \triangleright_C \text{ (ret } v, \Omega_2, \beta_2)\), so \((c_1; (c_2; c_3), \Omega) \triangleright_C \text{ (ret } v, \Omega_2, \beta_1, \beta_2)\).

(3) Suppose \((c_1, \Omega) \triangleright_C \text{ (cont, } \Omega_1, \beta_1)\), \((c_2, \Omega_1) \triangleright_C \text{ (cont, } \Omega_2, \beta_2)\) and \((c_3, \Omega_2) \triangleright_C \text{ (r, } \Omega_3, \beta_3)\). Then \((c_1; c_2, \Omega) \triangleright_C \text{ (cont, } \Omega_3, \beta_1, \beta_2)\), so \(((c_1; c_2); c_3, \Omega) \triangleright_C \text{ (r, } \Omega_3, \beta_1, \beta_2, \beta_3)\). It follows as well that \((c_2; c_3, \Omega_1) \triangleright_C \text{ (r, } \Omega_3, \beta_2, \beta_3)\), so \((c_1; (c_2; c_3), \Omega) \triangleright_C \text{ (r, } \Omega_3, \beta_1, \beta_2, \beta_3)\).

C Detailed Shallow Simulation

C.1 Shallow Semantics

Here are some rules not listed in the main content.

\[\begin{array}{c}
(c_1; \Omega) \triangleright_C \text{ (ret } e, \Omega_1, \alpha) \\
(c_2; \Omega) \triangleright_C \text{ (ret } e, \Omega_1, \alpha) \\
\text{if } e \text{ then } c_1 \text{ else } c_2, \Omega) \triangleright_C \text{ (r, } \Omega', h) \\
\text{while } e \text{ do } c, \Omega) \triangleright_C \text{ (r, } \Omega', h)
\end{array}\]

Here’s the dynamic semantics of the two new commands:

\[
\begin{array}{c}
\text{push, } \alpha \triangleright_C \text{ (cont, } \emptyset, \Omega, \langle\rangle) \\
\text{pop } \alpha = e, \cdot, \alpha \triangleright_C \text{ (cont, } \overline{\alpha} \mapsto \emptyset, \langle\rangle)
\end{array}
\]

As with the dynamic semantics, we can show that command concatenation is associative.

FACT 3.

\[S((c_1; c_2); c_3; \Omega) \triangleright_C \text{ (r, } \Omega', h) \iff (c_1; (c_2; c_3), \Omega) \triangleright_C \text{ (r, } \Omega', h)\]

The proof is similar to Fact 2.

We say an R-expression \(e\) (or L-expression \(e\)) is normalized in a symbolic stack \(\Omega\), denoted as \(\text{rnorm}(\Omega, e)\) (or \(\text{lnorm}(\Omega, e)\), if all its sub-expressions cannot be further simplified.

\[
\begin{array}{ll}
\text{1norm}(\Omega, e) & \iff \text{1norm}(\Omega, e) \land \text{rnorm}(\Omega, e) \\
\text{1norm}(\Omega, \{e\}) & \iff \text{1norm}(\Omega, e) \\
\text{1norm}(\Omega, \{e, l\}) & \iff \text{1norm}(\Omega, e)
\end{array}
\]

These definitions establish another helpful consequence.

FACT 4.

\[e\Omega_R = e' \implies \text{rnorm}(\Omega, e') \land [e]_L\Omega = e' \implies \text{lnorm}(\Omega, e').\]

FACT 5. No pointer dereference will appear in a generated program.

C.2 Proof of Theorem 1

Before the proof itself, we first present some simple lemmas about interacting with the actual stack and the symbolic stack, and then show two further lemmas about the correctness of simplifying R-expressions and L-expressions.

**Lemma 7.** If \([e]_L = \mu\), then for any stack \(\Omega\), \([e]_R\Omega = \mu\).

The proof follows from structural induction on L-expressions.

**Lemma 8.** \(\mathcal{R}(\text{load}(\Omega, \mu)) = \text{load}(\mathcal{R}(\Omega), \mu)\).

**Lemma 9.** \(\mathcal{R}(\text{store}(\Omega, \mu, v)) = \text{store}(\mathcal{R}(\Omega), \mu, \mathcal{R}(v))\).

**Lemma 10.** \(\mathcal{R}(\alpha \mapsto v) = \mathcal{R}(\Omega)[\alpha \mapsto \mathcal{R}(v)]\).

**Lemma 11.** \(\mathcal{R}(\omega \cdot \alpha) = \mathcal{R}(\omega) \cdot \mathcal{R}(\alpha)\).

**Lemma 12.** \(\text{load}(\Omega, \mu) = v \iff [\mathcal{R}(\omega), \mu]_R^{\Omega} = v\).

**Lemma 13.** If \(\text{rnorm}(\Omega, \alpha)\), then \([e]_R = \mathcal{R}(\mathcal{R}(\Omega))\).

**Lemma 14.** If \(\text{1norm}(\mathcal{R}(\Omega), \alpha)\), then either \(\alpha \notin \text{dom}[\cdot]_L\Omega\) or \([e]_L = \mathcal{R}(\mathcal{R}(\Omega))\).

The two lemmas can be proven together by structural induction on R-expressions and L-expressions. We discuss a few tricky cases here.

**R-expression e.** Suppose \([e]_R = v\), \([e]_L = \mu\) and \(v = \text{load}(\Omega, \mu)\).

**Lemma 16.** \([e]_L = \mu \iff \exists \alpha', \mathcal{R}(\mathcal{R}(\Omega)) = e' \land [e']_L = \mu\).

The goal by \([e']_R^{\Omega} = \text{load}(\Omega, \mathcal{R}(\what{\mu}) = \text{load}(\Omega, \mu) = v\).
Figure 18: Dynamic Semantics

Figure 19: Converting symbolic values and expressions.

(2) \( e' \in \text{dom}[-] \). Lemma 7 tells us \( [\cdot]_{L} = [\cdot]_{L}^{\Omega} = \mu \). So \( \text{load}(\mathcal{R}(\mu), e') = \text{load}(\mathcal{R}(\Omega), \mu) \), \( [\cdot]_{L}^{\Omega} = \mu \), \( \mu'_{R} \) which equals to \( \mathcal{R}(\Omega) \), \( \mu'_{R} \) according to lemma 8. Then we finish the goal using lemma 12: \( [\mathcal{R}(\mu), \mu]_{R}^{\Omega} = \mu \).

R-expression & \( e' \). Suppose \( \alpha \in \mathcal{R}(\Omega) = \text{Vref} \mu \) and \( [\cdot]_{L}^{\Omega} = \mu \). Lemma 16 shows there exists \( e' \) such that \( [\cdot]_{L}^{R(\Omega)} = \mu'_{R} \) and \( [\cdot]_{L}^{\Omega} = \mu'_{R} \). We want to prove \( [\cdot]_{L}^{R(\Omega)} = \mu'_{R} \) and \( [\cdot]_{L}^{\Omega} = \mu'_{R} \). The first goal is direct from the definition. As the second goal for definition, by definition \( [\cdot]_{L}^{R(\Omega)} = \text{Vref} \mu \).

L-expression \( e \). Suppose \( [\cdot]_{L}^{\Omega} = \text{Vindex} \mu \). \( [\cdot]_{L}^{\Omega} \) is \( \mu \) and \( [\cdot]_{L}^{\Omega} = \text{Vindex} \mu \). Lemma 15 and 16 guarantee \( e \) and \( e' \) can be simplified to \( e' \) and \( e'_{R} \), respectively, where \( [\cdot]_{L}^{R(\Omega)} = \mu'_{R} \), \( [\cdot]_{L}^{R(\Omega)} = \mu'_{R} \), \( [\cdot]_{L}^{\Omega} = \text{Vindex} \mu \). By definition we know \( [\cdot]_{L}^{R(\Omega)} = [\cdot]_{L}^{\Omega} \). Also \( [\cdot]_{L}^{R(\Omega)} = [\cdot]_{L}^{\Omega} \).

L-expression \( * e \). Suppose \( [\cdot]_{L}^{\Omega} = \text{Vref} \mu \). Lemma 16 shows there exists \( e' \) such that \( [\cdot]_{L}^{R(\Omega)} = e'_{R} \) and \( [\cdot]_{L}^{\Omega} = \text{Vref} \mu \).

Depending on whether \( e' \) can be simplified or not, there are two cases.

1. \( e' = e \) for some \( e \). Then \( [\cdot]_{L}^{R(\Omega)} = e \) by definition. We also know \( [\cdot]_{L}^{\Omega} = \mu \) because \( [\cdot]_{L}^{\Omega} = \text{Vref} \mu \), so this case is proven.

2. \( e' \) can not be simplified. Then \( [\cdot]_{L}^{R(\Omega)} = e \cdot e' \) by definition. Also \( [\cdot]_{L}^{\Omega} = \mu \) by definition.

We can finally prove theorem 1 by structural induction on commands.

\( \text{assert} \ e \). Suppose \( \text{assert} \ e \). Lemma 15 guarantees there exists \( e' \) such that \( [\cdot]_{L}^{R(\Omega)} = e' \) and \( [\cdot]_{L}^{\Omega} = \mu \). We now have two subgoals: \( \text{assert} \ e \) and \( \text{assert} \ e' \). Both can be proven using semantics definitions.

\( x = f(e_{1}, \ldots, e_{n}) \). Suppose \( x = f(e_{1}, \ldots, e_{n}) \). Lemma 15 guarantees there exist \( e' \) and \( e'_{R} \) such that \( [\cdot]_{L}^{R(\Omega)} = e' \) and \( [\cdot]_{L}^{\Omega} = \mu \). By definition we know \( [\cdot]_{L}^{R(\Omega)} = [\cdot]_{L}^{\Omega} \). Also \( [\cdot]_{L}^{R(\Omega)} = [\cdot]_{L}^{\Omega} \).

\( [\cdot]_{L}^{\Omega} = \text{Vref} \mu \). Lemma 16 shows there exists \( e' \) such that \( [\cdot]_{L}^{R(\Omega)} = e' \) and \( [\cdot]_{L}^{\Omega} = \text{Vref} \mu \).
We can prove the goal by concatenating these commands together.

\( \tau_x = \text{sstore} = \epsilon \alpha \) by subexpressions, respectively. We also define \( \text{reside in } \text{prims} \).

Due to Fact 5 we can ignore dereferencing pointers. Let the set \( \text{prims}(\mu) \) contain all smallest L-values with \( K_0 \) access level that reside in \( \mu \). We use \( \text{prims}(\mu) \) to define two functions \( \text{Lread}(\epsilon) \) and \( \text{Rread}(\epsilon) \) that collect all of the smallest \( K_0 \) L-values from \( \epsilon \) and \( \epsilon \)'s subexpressions, respectively. We also define \( \text{Pointed}(\epsilon) \) to collect all of the smallest L-values with \( K_0 \) access level that are pointed to by \( \epsilon \)'s subexpressions. It is used to analyze a sandboxed function’s arguments, as every pointer passed into a sandboxed function is potentially read. We omit its definition as it is near identical to \( \text{Rread}(\epsilon) \).

\[
\begin{align*}
\text{Lread}(\epsilon) &= \emptyset \\
\text{Lread}(\epsilon[\ell]) &= \text{Lread}(\epsilon) \cup \text{Rread}(\epsilon[\ell]) \\
\text{Lread}(\epsilon.J) &= \text{Lread}(\epsilon)
\end{align*}
\]

\[
\begin{align*}
\text{Rread}(\ell \ n) &= \emptyset \\
\text{Rread}(\& \ e) &= \emptyset \\
\text{Rread}(\epsilon) &= \text{prims}(\epsilon[\ell]) \cup \text{Lread}(\epsilon) \\
\text{Rread}(\ell \{e_1, \ldots\}) &= \cup \text{Rread}(e_i) \\
\text{Rread}(\ell \{e_0, \ldots\}) &= \cup \text{Rread}(\epsilon)
\end{align*}
\]

Ref. (\( \tau \alpha = e \)) = \text{Rread}(\epsilon)

\[
\begin{align*}
\text{DEF}(\tau \alpha = e) &= \text{prims}(\text{Vvar } \alpha) \\
\text{DEF}(\tau \alpha = f(e_1, \ldots)) &= \text{prims}(\text{Vvar }\alpha) \cup \text{Pointed}(\epsilon[\ell]) \\
\text{DEF}(\tau \alpha = f(e_1, \ldots) \text{ R S}_R \text{ W S}_W) &= \text{prims}(\text{Vvar }\alpha) \cup \text{S}_{\text{W}} \\
\text{DEF}(\text{assert } \epsilon) &= \emptyset \\
\text{DEF}(\epsilon = e) &= \text{prims}(\epsilon[\ell])
\end{align*}
\]

Some simple facts about the \( \approx \) relation.

\[
\begin{align*}
\text{FACT 6. } \Omega \approx_{\text{V}} \Omega' \land \Omega \approx_{\text{S}} \Omega' &\implies \Omega \approx_{\text{S}_1 \cup \text{S}_2} \Omega'. \\
\text{FACT 7. } \approx_{\text{S}} &\text{ is an equivalence relation.} \\
\text{FACT 8. } \Omega \approx_{\text{V}} \Omega' \land S_2 \subseteq S_1 &\implies \Omega \approx_{S_1} \Omega'.
\end{align*}
\]

This lemma says executing a command does not change values defined outside \( \text{DEF}(\epsilon) \).

\[
\begin{align*}
\text{Lemmn 17.} \forall \mu \not\in \text{DEF}(\epsilon), (\epsilon, \Omega) \rightarrow_C (\text{cont}, \Omega', \beta) &\implies \Omega \approx_{(\mu)} \Omega'
\end{align*}
\]

### D.1 Proof of Theorem 4

Here we restate the theorem:

If both \( (c_1; \ldots; c_{j-1}, \Omega) \rightarrow_C (\text{cont}, \Omega_{j-1}, \beta_{j-1}) \) and \( (c_j, \Omega_{j-1}) \rightarrow_C (\text{cont}, \Omega_j, \beta_j) \), then

\[
\begin{align*}
\forall \Omega'_{j-1}, \Omega_j \approx_{\text{DEF}(\neg \cdot)} \Omega'_{j-1} &\implies \\
\exists \Omega'_{j}, (c_j, \Omega_{j-1}) &\rightarrow_C (\text{cont}, \Omega'_j, \beta'_j) \land \Omega_j \approx_{(\text{DEF}(\neg \cdot))} \Omega'_{j}.
\end{align*}
\]

This theorem indicates two facts about the live variable analysis:

(1) The variables dependent by \( c_j \), i.e., \( \text{DEF}(\neg \cdot) \), are enough to execute \( c_j \), and

(2) When the variables in \( \text{DEF}(\neg \cdot) \) are the same in the starting stacks, after executing \( c_j \), the ending stacks are no different in terms of \( \text{DEF}(\neg \cdot) \).

These two facts allow a program chunk to safely execute with part of the stack rather than reconstructing the entire stack.

Proof. Firstly, using lemma 3, we know \( \Omega_{j-1} \) and \( \Omega'_{j-1} \) are similar up to \( \text{REF}(c_j) \). By the definition of \( \text{REF}(c_j) \), we know during executing \( c_j \), it only reads values from \( \text{REF}(c_j) \). Thus, the behavior is also determined by them. We do not show detailed proof here.

So there must exist an ending stack \( \Omega'_{j} \) where \( c_j \) can reach from \( \Omega'_{j-1} \) and \( \Omega_j \approx_{\text{DEF}(c_j)} \Omega'_j \). Now using lemma 17, we know \( \Omega_{j-1} \approx_{\text{REF}(c_j) \rightarrow \text{DEF}(c_j)} \Omega_j \) and \( \Omega'_{j-1} \approx_{\text{REF}(c_j) \rightarrow \text{DEF}(c_j)} \Omega'_{j} \). Since
\( \Omega_{j-1} \equiv_{\text{Ref}(c_j)} \Omega'_{j-1} \), we can use facts 7 and 8 to derive \( \Omega_j \equiv_{\text{Ref}(c_j)} \Omega'_{j-1} \).

Finally, fact 6 allows us to combine the similarity between \( \Omega_j \) and \( \Omega'_{j} \) in terms of \( \text{Ref}(c_j) \) and \( \text{Ref}(c_j-\text{DEF}(c_j)) \) into \( \Omega_j \equiv_{\text{Ref}(c_j) \cup \text{DEF}(c_j)} \Omega'_{j} \), then use lemma 2 we can prove \( \Omega_j \equiv_{\text{Def} (i,\ldots, j)} \Omega'_{j} \).

**Theorem 18.** Using Theorem 4’s setting, if \( i < j < k \), then \( \Omega_{j-1} \equiv_{\text{Def}(i,k)} \Omega_j \).

Proof. We will prove it using lemma 17 by showing each variable in \( \text{DEF}(i,k) \) does not occur in \( \text{DEF}(c_j) \), i.e., \( \text{DEF}(c_j) \cap \text{DEF}(i,k) = \emptyset \). By definition, if \( \mu \in \text{DEF}(i,k) \), then \( \mu \mapsto k \in \text{LIVE}(i) \). We then use the following lemma to prove \( \mu \notin \text{DEF}(c_j) \).

**Lemma 19.** \( \mu \mapsto k \in \text{LIVE}(i) \implies \forall j \in [i,k), \mu \mapsto k \in \text{LIVE}(j) \).

This lemma is direct from the definition of \( \text{LIVE}(\cdot) \): a dependency \( \mu \mapsto k \) is added when analyzing \( c_k \) and will stay in the live set until \( \mu \) is used in a command. Since it still exists in \( \text{LIVE}(i) \), all live sets between \( i \) and \( k \) must not resolve it.

Returning to the theorem, we know \( \mu \mapsto k \in \text{LIVE}(j-1) \) using this lemma. So by definition of \( \text{LIVE}(j-1) \), \( \mu \notin \text{DEF}(c_j) \).

## D.2 Compromise in Analysis

Note this section’s discussion assumes the reader has read Sec. 5.2. One part of the efficient-cutting problem’s goal is to minimize the sum of all chunk’s external dependencies. However, in many cases, the cutting cost could be smaller. Consider

\[
\text{suppose the cutting cost of one } \text{pvt int } x \text{ value is N. If Lian decides to partition L1 into one chunk, L2 and L3 into another chunk, then the OPB's encoding considers the cutting cost to be 2N. However, since } x \text{ is not updated between L2 and L3, L3 can use } x \text{'s local value without synchronization.}
\]

We could have used a more accurate encoding, but that will significantly increase the complexity and slow down the solver. So here we make a compromise by moving the dependency between L1 and L3 to L2 and L3. Now the dependency becomes local, so it no longer contributes to the total cutting cost.

We can rigorously prove in most cases, this approach does not reduce the total cutting cost in the encoding: If the program is cut into \((1,2,3)\), \((1,2),(3)\) or \((1,2),(3)\), the total cost does not change (even if there are other commands involved). There is only one case that results in bigger cutting cost: \((1,3),(2)\). In other words, we will overestimate the total cost. This may not be a bad thing as now our encoding is prone to cut the program without breaking its original encoding. Also, if the program has better performance by cutting 2 out, then usually the user can rewrite the program by moving commands that have close data dependencies together.

### E Detailed Deep Simulation

Here are some unlisted semantics rules.

\[
d\text{load}(\Omega, e) = \begin{cases} \text{load}(\Omega, \mu), \mu | \Omega \text{ when } e | J_{L_1} = \mu \\ \emptyset \text{ otherwise} \end{cases}
\]

\[
d\text{store}(\Omega, e, e) = \begin{cases} \text{store}(\Omega, \mu, [e]_{R_1}) \text{ when } e | J_{L_1} = \mu \\ \emptyset \text{ otherwise} \end{cases}
\]

\[
\text{new}(x) = \alpha \quad [e]_{R_1} = e'
\]

\[
(\tau x = e, \Omega) \iff_{C} (\text{cont, } \Omega(x \mapsto [e]_{R_1}))
\]

\[
[e]_{L_1}^\Omega = e'
\]

\[
(\text{e, } \Omega) \iff_{C} (\text{cont, dstore}(\Omega, e', e'))
\]

\[
(\text{return e, } \Omega) \iff_{C} (\text{cont, dstore}(\Omega, e', e'))
\]

\[
(\text{assert e, } \Omega) \iff_{C} (\text{cont, } \Omega)
\]

\[
(c_1, \Omega) \iff_{C} (\text{cont, } \Omega) \quad (c_2, \Omega_1) \iff_{C} (\text{cont, } \Omega)
\]

\[
(c_1, \Omega) \iff_{C} (c_1; c_2, \Omega) \iff_{C} (c_2, \Omega) \iff_{C} (c, \Omega_2)
\]

\[
[e]_{R_1}^\Omega = \ell \text{ true} \quad (c_1, \Omega) \iff_{C} (c, \Omega)
\]

\[
(\text{if e then c1 else c2, } \Omega) \iff_{C} (c_1, \Omega) \iff_{C} (c_2, \Omega)
\]

\[
[e]_{R_1}^\Omega = \ell \text{ false} \quad (c_2, \Omega) \iff_{C} (c, \Omega)
\]

\[
(\text{if e then c1 else c2, } \Omega) \iff_{C} (c_2, \Omega) \iff_{C} (c, \Omega)
\]

\[
[c; \text{while do c, } \Omega] \iff_{C} (c, \Omega)
\]

\[
[\text{while do c, } \Omega] \iff_{C} (c, \Omega)
\]

\[
(\text{new}(x) = \alpha \quad \eta \in \{\text{normal, atomic, plocal}, \text{box}\})
\]

\[
f \mapsto \text{internal}(\eta, \tau_{x_1, x_2, \ldots, x_n} \mapsto c, \eta) \eta \in \Lambda^\eta
\]

\[
\forall \eta \in [1,n], [e]_{R_1}^\Omega = e' \text{ and new}(x) = d_{i1}
\]

\[
(c, \text{while do c, } \Omega) \iff_{C} (\text{cont, } \Omega(\alpha \mapsto [e]_{R_1}))
\]

\[
(\tau x = f(x_1, \ldots, x_n), \Omega) \iff_{C} (\text{cont, } \Omega(\alpha \mapsto [e]_{R_1}))
\]
Theorem 6 can be derived from the following lemma.

**Lemma 22.** \( \forall j \in \text{chunk}^{-1}(t) \), \( \Omega^\pre_j \cong S \Omega_{j-1} \) where \( S = \phi^t_{j-1} \lor \text{DEP}(\neg, j) \).

We prove it together with the following lemma by induction on the sequential program.

**Lemma 23.** \( \forall j \in [1, n], \Omega^\post_j \cong S \Omega_{j-1} \) where \( S = \phi^t_j \).

Subgoal 1. Suppose \( \text{chunk}(j) = t \) and
\[ \Omega^\post_{j-1} \cong S \Omega_{j-1} \) where \( S = \phi^t_{j-1} \),
we want to derive
\[ \Omega^\pre_j \cong S' \Omega_{j-1} \) where \( S' = \phi^t_{j-1} \lor \text{DEP}(j, \neg) \).

We know \( \Omega^\post_{j-1} \cong S \Omega^\pre_j \) as \( \text{sync}_j \) only updates external dependencies, so fact 7 ensures \( \Omega^\pre_j \cong S \Omega_{j-1} \). Also, theorem 5 guarantees the correctness of \( \text{sync}_j \), i.e.,
\[ \Omega^\pre_j \cong S' \Omega_{j-1} \) where \( S'' = \bigcup_{\text{chunk}(i) \neq t} \text{DEP}(i, j) \).

Finally, we can prove
\[ \phi^t_{j+1} \lor \bigcup_{\text{chunk}(i) \neq t} \text{DEP}(i, j) \subseteq \phi^t_{j-1} \lor \text{DEP}(\neg, j) \),
then use fact 6 and 8 to derive
\[ \Omega^\pre_j \cong S \Omega_{j-1} \).

Subgoal 2. Suppose \( \text{chunk}(j) = t \) and
\[ \Omega^\pre_j \cong S' \Omega_{j-1} \) where \( S' = \phi^t_{j-1} \lor \text{DEP}(\neg, j) \),
we want to derive
\[ \Omega^\post_j \cong S \Omega_j \) where \( S = \phi^t_j \).

We divide \( \phi^t_j \) into two parts \( S_1 \) and \( S_2 \) where \( S_1 = \bigcup_{\text{chunk}(k) = t} \text{DEP}(j, k) \) contains all internal dependencies coming out of \( j \), and \( S_2 = \phi^t_j - S_1 \) contains all internal dependencies coming before \( j \). Theorem 4 ensures \( \Omega^\pre_j \cong S_1 \Omega_j \), and theorem 18 ensures \( \Omega_{j-1} \cong S_2 \Omega_j \) and \( \Omega^\pre_j \cong S_2 \Omega^\post_j \). Notice \( S_2 \subseteq \phi^t_{j-1} \), so we can combine with the

**Figure 20:** The effect of adding atomic annotations to the MT program for 16-block inputs.

**Figure 21:** The effect of adding atomic annotations to the MT program for 64-block inputs.

**Figure 22:** The effect of adding atomic annotations to the MT program for 8-block inputs.
(a) Program structure of GD with `update_all()` annotated with atomic keyword

(b) Program structure of GD without `update_all()` being atomic

(a) Program structure of MT with `sha256_node()` annotated with atomic keyword

(b) Program structure of MT without `sha256_node()` being atomic