Thesis on

Counterexample-Guided Verification of Imperative Programs Against Implementation Agnostic Functional Specification

by

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(2020CSY7569)

Under the guidance of

Prof. Sorav Bansal
(Computer Science and Engineering)

Submitted in the partial fulfillment
of the requirements for the degree of

Master of Science (Research)

to the

Department of Computer Science and Engineering
Indian Institute of Technology Delhi

June 2023

Certificate

This is to certify that the thesis titled **"Counterexample-Guided Verification of Imperative Programs Against Implementation Agnostic Functional Specification"**, being submitted by **Mr.Indrajit Banerjee**, to the Indian Institute of Technology, Delhi, for award of the degree **Master of Science (Research)**, is a bona fide record of the research work done by him under my supervision. The contents of this thesis, in full or in parts, have not been submitted to any other Institute or University for the award of any degree or diploma.

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Acknowledgments

I would like to sincerely thank my thesis supervisor Prof. Sorav Bansal for his continuous support during my study and research. His guidance, patience, motivation and long discussions provided a strong platform with clear visibility and research direction.

Besides my advisor, I would like to thank the following members of my Student Research Committee for their insightful comments and encouragement that helped me to widen my research from various perspectives:
Prof. Sanjiva Prasad (Dept. of CSE, IIT Delhi)
Prof. Kumar Madhukar (Dept. of CSE, IIT Delhi)
Mr. Akash Lal (Microsoft Research Lab, India)

I am grateful to our research group members: Abhishek Rose, Shubhani at IIT Delhi for their help and motivating discussions on various topics related to my research.

Indrajit Banerjee

Abstract

We describe an algorithm capable of checking equivalence of two programs that manipulate recursive data structures such as linked lists, strings, trees and matrices. The first program, called specification, is written in a succinct and safe functional language with algebraic data types (ADT). The second program, called implementation, is written in C using arrays and pointers. Our algorithm, based on prior work on counterexample guided equivalence checking, automatically searches for a sound equivalence proof between the two programs.

We formulate an algorithm for discharging proof obligations containing relations between recursive data structure values across the two diverse syntaxes, which forms our first contribution. Our proof discharge algorithm is capable of generating falsifying counterexamples in case of a proof failure. These counterexamples help guide the search for a sound equivalence proof and aid in inference of invariants. As part of our proof discharge algorithm, we formulate a program representation of values. This allows us to reformulate proof obligations due to the top-level equivalence check into smaller nested equivalence checks. Based on this algorithm, we implement an automatic (push-button) equivalence checker tool named S2C, which forms our second contribution.

S2C is evaluated on implementations of common string library functions taken from popular C library implementations, as well as implementations of common list, tree and matrix programs. These implementations differ in data layout of recursive data structures as well as algorithmic strategies. We demonstrate that S2C is able to establish equivalence between a single specification and its diverse C implementations.

Keywords: *Equivalence checking; Bisimulation; Recursive Data Structures; Algebraic Data Types;*

Contents

1 Introduction

The problem of equivalence checking between a functional specification and an implementation written in a low level imperative language such as C has been of major research interest and has several important applications such as (a) program verification, where the equivalence checker is used to verify that the C implementation behaves according to the specification and (b) translation validation, where the equivalence checker attempts to generate a proof of equivalence across the transformations (and translations) performed by an optimizing compiler and more.

The verification of a C implementation against its manually written functional specification through manually-coded refinement proofs has been performed extensively in the seL4 microkernel [25]. Frameworks for program equivalence proofs have been developed in interactive theorem provers like Coq [16] where correlations and invariants are manually identified during proof codification. On the other hand, programming languages like Dafny [27] offer automated program reasoning for imperative languages with abstract data types such as sets and arrays. Such languages perform automatic compile-time checks for manually-specified correctness predicates through SMT solvers. Additionally, there exists significant prior work on translation validation [32, 42, 39, 41, 26, 44, 45, 36, 43, 28, 24, 29, 11, 38, 15, 22, 37, 31] across low level programming languages such as C and assembly. In most of these applications, soundness in critial, i.e., if the equivalence checker determines the programs to be equivalent, then the programs are indeed equivalent and evidently has equivalent observable behaviour. On the other hand, a sound equivalence checker may be incomplete and fail to prove the programs to be equivalent, even if they were equivalent.

We present S2C, a *sound* algorithm to automatically (push-button) search for a proof of equivalence between a functional specification (written in Spec) and its optimized C implementation. We will demonstrate how S2C is capable of proving equivalence of multiple equivalent C implementations with vastly different (a) data layouts (e.g. array, linked list representations of a *list*) and (b) algorithmic strategies (e.g. alternate algorithms, optimizations) against a *single* functional specification. This opens the possibility of regression verification [40, 20], where S2C can be used to automate verification across software updates that change

memory layouts for data structures.

1.1 Summary

We restrict our attention to programs that construct, read, and write to recursive data structures. In languages like C, pointer and array based implementations of these data-structures are prone to safety and liveness bugs. Similar recursive data structures are also available in safer functional languages like Haskell, where algebraic data types (ADTs) [13] ensure several safety properties. We define a minimal functional language, called Spec, that enables the safe and succinct specification of programs manipulating and traversing recursive data structures. Spec is equipped with ADTs as well as boolean and bitvector (`i<N>`) types.

Next, we give a brief overview of our approach through an example. This allows us to introduce the major subgoals and we state our primary contributions in the next section.

```
A0: type List = LNil | LCons (val:i32, tail:List).
A1:
A2: fn mk_list_impl (n:i32) (i:i32) (l:List) : List =
A3:    if i ≥u n then l
A4:              else make_list_impl(n, i+1i32, LCons(i, l)).
A5:
A6: fn mk_list (n:i32) : List = mk_list_impl(n, 0i32, LNil).
```

(a) Spec Program

```
B0:  typedef struct lnode {
B1:    unsigned val; struct lnode* next; } lnode;
B2:
B3:  lnode* mk_list(unsigned n) {
B4:    lnode* l = NULL;
B5:    for (unsigned i = 0; i < n; ++i) {
B6:      lnode* p = malloc(sizeof lnode);
B7:      p→val = i; p→next = l; l = p;
B8:    }
B9:    return l;
B10: }
```

(b) C Program with `malloc()`

Figure 1: Spec and C Programs constructing a Linked List.

```
S0: List mk_list (i32 n) {
S1:   List l := LNil;
S2:   i32  i := 0_i32;
S3:   while ¬(i ≥_u n):
S4:     l := LCons(i, l);
S5:     i := i + 1_i32;
S6:   return l;
SE: }
```

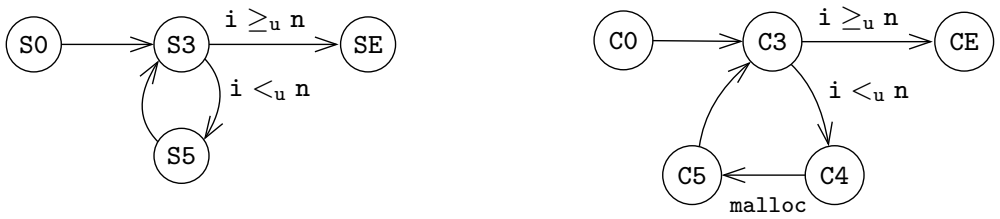
(a) (Abstracted) Spec IR

```
C0: i32 mk_list (i32 n) {
C1:   i32 l := 0_i32;
C2:   i32 i := 0_i32;
C3:   while i <_u n:
C4:     i32 p := malloc_C4(sizeof lnode);
C5:     m := m[&(p ->^m_lnode val) ← i]_i32;
C6:     m := m[&(p ->^m_lnode next) ← l]_i32;
C7:     l := p;
C8:     i := i + 1_i32;
C9:   return l;
CE: }
```

(b) (Abstracted) C IR

Figure 2: IRs for the Spec and C Programs in figs. 1a and 1b respectively.

Figures 1a and 1b show the construction of lists in Spec and C respectively. The `List` ADT in the Spec program in defined at line `A0` in fig. 1a. An empty `List` is represented by the constructor `LNil`, where as a non-empty list uses the `LCons` constructor to combine its first value (`val` :`i32`) and the remaining list (`tail` :`List`). The inputs to a Spec procedure are its well-typed arguments, which may include recursive data structure values. The inputs to a C procedure are its explicit arguments and the implicit state of program memory at procedure entry. We lower both Spec and C programs to a common intermediate representation (IR) as shown in figs. 2a and 2b. For the Spec program in fig. 1a, the tail-recursive function `mk_list_impl` is converted to a loop and inlined in the top-level function `mk_list`. For the C program in fig. 1b, the sizes and memory layouts of both scalar (e.g., `unsigned`) and compound (e.g., `struct lnode`) types are concretized in the IR.



(a) CFG of Spec Program

(b) CFG of C Program

Figure 3: CFG representation for Spec and C IRs shown in figs. 2a and 2b

Figures 10a and 10b show the Control-Flow Graph (CFG) representation of the Spec and C IR programs in figs. 2a and 2b respectively. Each node represent a PC

Table 1: Node Invariants for Product-CFG in fig. 11

PC-Pair	Invariants
(S0:C0)	(P) $\mathtt{n}_S = \mathtt{n}_C$
(S3:C3)	(I1) $\mathtt{n}_S = \mathtt{n}_C$ (I2) $\mathtt{i}_S = \mathtt{i}_C$ (I3) $\mathtt{i}_S \leq_u \mathtt{n}_S$ (I4) $\mathtt{l}_S \sim \mathtt{Clist}_{\mathbb{m}}^{\mathtt{lnode}}(\mathtt{l}_C)$
(S3:C4) (S3:C5)	(I5) $\mathtt{n}_S = \mathtt{n}_C$ (I6) $\mathtt{i}_S = \mathtt{i}_C$ (I7) $\mathtt{i}_S <_u \mathtt{n}_S$ (I8) $\mathtt{l}_S \sim \mathtt{Clist}_{\mathbb{m}}^{\mathtt{lnode}}(\mathtt{l}_C)$
(SE:CE)	(E) $\mathtt{ret}_S \sim \mathtt{Clist}_{\mathbb{m}}^{\mathtt{lnode}}(\mathtt{ret}_C)$

location of its corresponding program, and each edge represent conditional transition between PCs through instruction execution. For brevity, we often represent a sequence of instructions with a single edge, e.g., in fig. 10b, the edge C5→C3 represents the path C5→C6→C7→C8→C3.

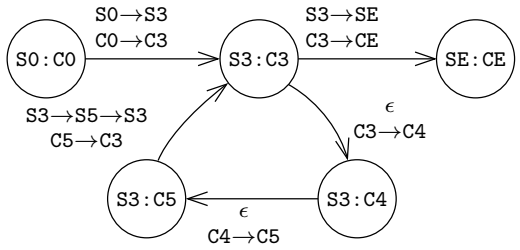


Figure 4: Product-CFG between the CFGs in figs. 10a and 10b

We construct a *bisimulation relation* to identify equivalence between the two programs. A bisimulation relation correlates the transitions of Spec and C programs in lockstep, such that the lockstep execution ensures identical observable behavior. A bisimulation relation between two programs can be represented using a *product program* [43] and the CFG representation of a product program is called a *product-CFG*. Figure 11 shows a product-CFG, that encodes the lockstep execution (bisimulation relation) between the CFGs in figs. 10a and 10b.

At each node of the product-CFG, invariants relate the states of the Spec and C program respectively. Table 2 lists invariants for the product-CFG in fig. 11. At the start node S0:C0 of the product-CFG, the precondition *Pre* (labeled (P)) ensures equality of input arguments $\mathtt{n}_S$ and $\mathtt{n}_C$ at the programs' entry. Inductive invariants (labeled (I)) are inferred at each intermediate product-CFG node (e.g., S3:C3) relating both programs' states. For example, at node S3:C5, (I6) $\mathtt{i}_S = \mathtt{i}_C$ is an inductive invariant.

In table 2, the invariant (I4) $\mathtt{l}_S \sim \mathtt{Clist}_{\mathbb{m}}^{\mathtt{lnode}}(\mathtt{l}_C)$ is an example of a recursive

relation and represents equality between the Spec `List` variable $\mathtt{l}_S$ and the `List` represented by chasing the `lnode` pointers starting at $\mathtt{l}_C$. $\mathtt{Clist}^{\mathtt{lnode}}_{\mathtt{m}}$ is an example of a *lifting constructor* that 'lifts' a C pointer value (pointing to an object of type `struct lnode`) and the C memory state `m` to a Spec `List` value, and is defined as follows:

$$\begin{aligned} U_C : \mathtt{Clist}^{\mathtt{lnode}}_{\mathtt{m}}(p{:}\mathtt{i32}) = \underline{\mathtt{if}}\ p = 0\ \underline{\mathtt{then}}\ \mathtt{LNil} \\ \underline{\mathtt{else}}\ \mathtt{LCons}(p \xrightarrow{\mathtt{m}}_{\mathtt{lnode}} \mathtt{val}, \mathtt{Clist}^{\mathtt{lnode}}_{\mathtt{m}}(p \xrightarrow{\mathtt{m}}_{\mathtt{lnode}} \mathtt{next})) \end{aligned} \tag{1}$$

Product-CFG invariants involving recursive relations (e.g., (I4)) allow us to express equality between native Spec values with the C program state. Assuming that the precondition *Pre* ((P)) holds at the entry node `S0:C0`, a bisimulation check involves checking that the inductive invariants hold too, and consequently the postcondition *Post* ((E)) holds at the exit node `SE:CE`. Checking whether an invariant holds results in proof queries. These proof obligations are expressed as relational Hoare triples [12, 23] and discharged through a proof discharge algorithm i.e. a solver. We give a more formal exposition of the concepts introduced in this summary in the coming chapters.

1.2 Our Contributions

As previously summarized in section 1.1, showing equivalence of a Spec and a C program through a bisimulation proof requires three major procedures: (1) An algorithm for construction of a product-CFG by correlating program executions across the Spec and C programs respectively. (2) An algorithm for identification of inductive invariants at intermediate correlated PCs. (3) An algorithm for solving proof obligations containing recursive relations. Our major contributions are as follows:

- Proof Discharge Algorithm: Solving proof obligations ((3)) involving recursive relations is rather interesting and forms our primary contribution. We describe a *sound* proof discharge algorithm capable of tackling proof obligations involving recursive relations using off-the-shelf SMT solvers. Our proof discharge algorithm is also capable of reconstruction of counterexamples for the original proof query from models returned by the individual SMT

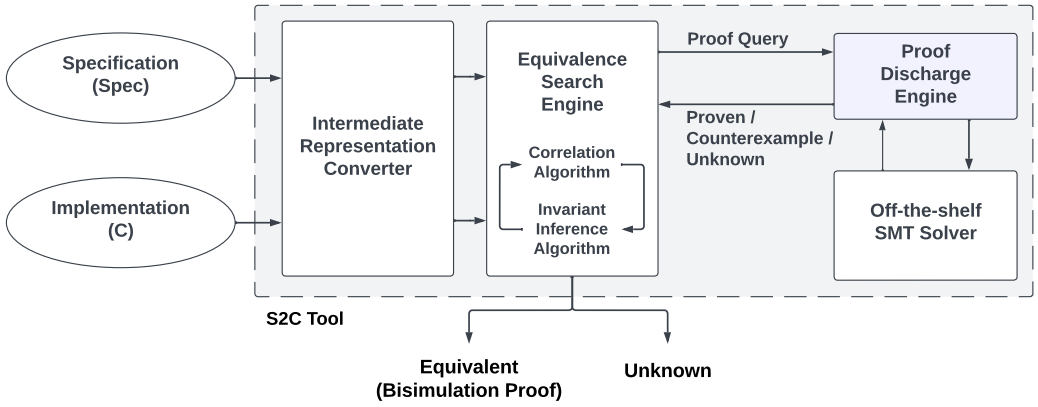


Figure 5: A brief overview of our equivalence checker algorithm S2C. The inputs to the algorithm are the Spec and C programs. S2C either successfully finds a bisimulation proof implying equivalence or soundly fails with an unknown verdict.

queries. These counterexamples are the backbone of counterexample-guided algorithms for ① and ② steps. As part of our proof discharge procedure, we reformulate equality of values (i.e. recursive relations) as equivalence of their corresponding programs and discharge these proof queries using a nested (albeit much simpler) bisimulation check.

- Spec-to-C Automatic Equivalence Checker Tool: Our second contribution is S2C, an equivalence checker tool capable of proving equivalence between a Spec and a C program automatically. S2C is based on the Counter tool[22] and uses modified versions of (a) counterexample-guided correlation algorithm for incremental construction of a product-CFG and (b) counterexample-guided invariant inference algorithm for inference of inductive invariants at correlated PCs in the (partially constructed) product-CFG. S2C discharges required verification conditions (i.e. proof obligations) using our Proof Discharge Algorithm. Figure 5 gives an overview of the complete algorithm.

2 Languages and Equivalence

2.1 The Spec Language

We start with a discussion on the Spec language. Spec supports recursive algebraic data types (ADT) similar to the ones available in most functional languages. Additionally, Spec is equipped with the following scalar types: `Unit`, Boolean (`Bool`) and Bitvector of length N (`i<N>`). ADTs can be thought of as 'sum of product' types where each constructor represents a variant and the arguments to each constructor represents its fields. Evidently, types in Spec can be represented in *first order recursive types* with `Product` and `Sum` type constructors and `Unit, Bool, i<N>` types (i.e., nullary type constructors) as follows:

$$T \rightarrow \mu\alpha.\ T \mid \texttt{Product}(T, \ldots, T) \mid \texttt{Sum}(T, \ldots, T) \mid \texttt{Unit} \mid \texttt{Bool} \mid \texttt{i}\langle\texttt{N}\rangle \mid \alpha$$

For example, the `List` type can be written as $\mu\alpha.Sum(Unit, Product(i32, \alpha))$.

The language also borrows its expression grammar heavily from functional languages. This includes the usual constructs like `let-in`, `if-then-else`, function application and the `match` statement for pattern-matching (i.e. deconstructing) sum and product values. Unlike functional languages, Spec only supports first order functions. Also, Spec does not support partial function application. Hence, we constrain our attention to C programs containing only first order functions. Spec is equipped with a special `assuming-do` construct for explicitly providing assertions. Spec also provides the typical boolean and bitvector operators for expressing computation in C succintly yet explicitly. This includes logical operators (e.g., `and`), bitvector arithmatic operators (e.g., `bvadd(+)`) and relational operators for comparing bitvectors interpreted as signed or unsigned integers (e.g., $\leq_{u,s}$).

$$
\begin{array}{rcl}
\langle \text{expr} \rangle & \rightarrow & \texttt{if}\ \langle \text{expr} \rangle\ \texttt{then}\ \langle \text{expr} \rangle\ \texttt{else}\ \langle \text{expr} \rangle \\
& | & \texttt{let}\ \langle \text{id} \rangle\ \texttt{=}\ \langle \text{expr} \rangle\ \texttt{in}\ \langle \text{expr} \rangle \\
& | & \texttt{match}\ \langle \text{expr} \rangle\ \texttt{with}\ \langle \text{match-clause-list} \rangle \\
& | & \texttt{assuming}\ \langle \text{expr} \rangle\ \texttt{do}\ \langle \text{expr} \rangle \\
& | & \langle \text{id} \rangle\ \texttt{(}\ \langle \text{expr-list} \rangle\ \texttt{)} \\
& | & \langle \text{data-cons} \rangle\ \texttt{(}\ \langle \text{expr-list} \rangle\ \texttt{)} \\
& | & \langle \text{expr} \rangle\ \texttt{is}\ \langle \text{data-cons} \rangle \\
& | & \langle \text{expr} \rangle\ \langle \text{scalar-op} \rangle\ \langle \text{expr} \rangle \\
& | & \langle \text{literal}_{\text{Unit}} \rangle \mid \langle \text{literal}_{\text{Bool}} \rangle \mid \langle \text{literal}_{\text{i<N>}} \rangle \\
\\
\langle \text{match-clause-list} \rangle & \rightarrow & \langle \text{match-clause} \rangle^* \\
\langle \text{match-clause} \rangle & \rightarrow & \mid \langle \text{data-cons} \rangle\ \texttt{(}\ \langle \text{id-list} \rangle\ \texttt{)} \Rightarrow \langle \text{expr} \rangle \\
\langle \text{expr-list} \rangle & \rightarrow & \epsilon \mid \langle \text{expr} \rangle\ \texttt{,}\ \langle \text{expr-list} \rangle \\
\langle \text{id-list} \rangle & \rightarrow & \epsilon \mid \langle \text{id} \rangle\ \texttt{,}\ \langle \text{id-list} \rangle \\
\\
\langle \text{literal}_{\text{Unit}} \rangle & \rightarrow & \texttt{()} \\
\langle \text{literal}_{\text{Bool}} \rangle & \rightarrow & \texttt{false} \mid \texttt{true} \\
\langle \text{literal}_{\text{i<N>}} \rangle & \rightarrow & [\texttt{0}\ldots\texttt{2}^{\text{N}}\texttt{-1}]
\end{array}
$$

Figure 6: Simplified expression grammar of Spec language

2.2 Intermediate Representations

As summarized in section 1.1, we lower both Spec and C programs to a common intermediate representation (IR) for comparison. IR is a Three-Address-Code (3AC) style intermediate representation. We often omit intermediate registers in the IR for brevity and ease of exposition, and refer to this as the *abstracted* IR.

Figures 12a and 12b show Spec and C programs that traverse a linked list and return the sum of all the values in the linked list. The corresponding IR programs are shown in figs. 8a and 8b.

During conversion of a Spec source (figs. 1a and 12a resp.) to IR (figs. 2a and 8a resp.), (a) `match` statements are lowered to explicit `if`-`then`-`else` conditionals where each branch represents a distinct constructor, (b) all tail-recursive calls are converted to loops while non-tail calls are preserved and (c) all helper functions are inlined at their call-site.

Similarly, the following is performed during conversion of a C source (figs. 1b and 12b resp.) to IR (figs. 2b and 8b resp.): (a) the sizes and memory layouts of both scalar (e.g., `unsigned`) and compound (e.g., `struct lnode`) types are

```
A0: type List = LNil | LCons (val:i32, tail:List).
A1:
A2: fn sum_list_impl (l:List) (sum:i32) : i32 =
A3:    match l with
A4:    | LNil => sum
A5:    | LCons(x, rest) => sum_list_impl(rest, sum + x).
A6:
A7: fn sum_list (l:List) : i32 = sum_list_impl(l, 0i32).
```

(a) Spec Program

```
B0:  typedef struct lnode {
B1:    unsigned val; struct lnode* next; } lnode;
B2:
B3:  unsigned sum_list(lnode* l) {
B4:    unsigned sum = 0;
B5:    while (l) {
B6:      sum += l→val;
B7:      l = l→next;
B8:    }
B9:    return sum;
B10: }
```

(b) C Program

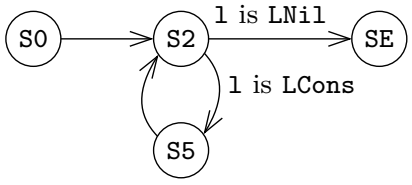
Figure 7: Spec and C Programs traversing a Linked List.

```
S0: i32 sum_list (List l) {
S1:   i32 sum := 0i32;
S2:   while ¬(l is LNil):
S3:     // (l is LCons);
S4:     sum := sum + l.val;
S5:     l   := l.next;
S6:   return sum;
SE: }
```

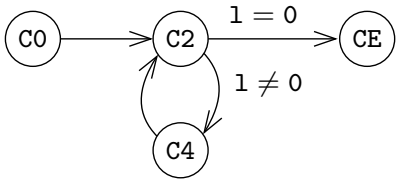
(a) (Abstracted) Spec IR

```
C0: i32 sum_list (i32 l) {
C1:   i32 sum := 0i32;
C2:   while l ≠ 0i32:
C3:     sum := sum + l →m_lnode val;
C4:     l   := l →m_lnode next;
C5:   return sum;
CE: }
```

(b) (Abstracted) C IR



(c) CFG of Spec Program



(d) CFG of C Program

Figure 8: IRs and CFGs of the Spec and C Programs in figs. 12a and 12b respectively.

concretized, (b) the program memory along with reads and writes to it are made explicit and (c) we annotate `malloc` calls with the call-site i.e. IR PC (e.g., $\texttt{malloc}_{\texttt{C4}}$ in fig. 2b).

The IR supports both scalar and ADT types available in Spec. Each ADT value is modeled as a key-value dictionary that maps each of its field names to the constituent values. These key-value pairs are accessed using the *accessor*-operator, e.g., $l.\texttt{val}$ and $l.\texttt{next}$ represents the first and second fields of the `LCons` constructor in fig. 8a. The IR also allows querying the top-level value constructor of an ADT value using the *is*-operator, e.g., l is `LNil` in fig. 8a. Importantly, $l.val$ is only well-formed if l is `LCons`. The construction of the Spec IR ensures the well-formedness of all expressions. Using the *accessor*- and *is*-operators, a `List` value l can be expanded as:

$$U_S : l = \underline{\texttt{if}}\ l\ \text{is}\ \texttt{LNil}\ \underline{\texttt{then}}\ \texttt{LNil}\ \underline{\texttt{else}}\ \texttt{LCons}(l.\texttt{val}, l.\texttt{next}) \tag{2}$$

In this expanded representation of l, the *sum-deconstruction* operator '$\underline{\texttt{if}}$-$\underline{\texttt{then}}$-$\underline{\texttt{else}}$' [1] conditionally deconstructs the sum type into its variants `LNil` and `LCons`. Equation (8) is called the *unrolling procedure* for the `List` variable l. We can similarly define the unrolling procedure for any ADT variable.

Pointers are converted to bitvectors and the C memory is modeled as a byte-addressable array $\mathtt{m}$ in the IR. Memory reads are represented using the following two C-like syntaxes: (a) "$p \xrightarrow{\mathtt{m}}_{\mathtt{T}} \mathtt{f}$" is equivalent to "*(`typeof(T.f)`*)(&$\mathtt{m}$[p+`offsetof(T,f)`])" i.e., it returns the bytes in the memory array $\mathtt{m}$ starting at address 'p+`offsetof(T,f)`' and interpreted as an object of type '`typeof(T.f)`' and (b) "$p[i]^{\mathtt{T}}_{\mathtt{m}}$" is equivalent to "*(T*)(&$\mathtt{m}$[p + i×`sizeof(T)`])" i.e., it returns the bytes in the memory array $\mathtt{m}$ starting at address '$p+i\times$`sizeof(T)`' and interpreted as an object of type '`T`'. "$\mathtt{m}[a \leftarrow v]_{\mathtt{T}}$" represents an array that is equal to $\mathtt{m}$ everywhere except at addresses [a, a+`sizeof(T)`) which contains the value v of type '`T`'.

Figures 8c and 8d show the Control-Flow Graph (CFG) representation of the Spec and C IRs in figs. 8a and 8b respectively. Each CFG node represents a IR

[1] The sum-deconstruction operator '$\underline{\texttt{if}}$-$\underline{\texttt{then}}$-$\underline{\texttt{else}}$' for an ADT T must contain exactly one branch for each value constructor of T. For example, '$\underline{\texttt{if}}$-$\underline{\texttt{then}}$-$\underline{\texttt{else}}$' for the `List` type must have exactly two branches of the form `LNil` and $\texttt{LCons}(e_1, e_2)$ for some expressions e_1 and e_2.

PC location of the program and edges represent transitions through execution of instructions. Each edge is associated with: (a) a *edge condition* (the condition under which that edge is taken), (b) a *transfer function* (how the program state is mutated if that edge is taken) and (c) a *UB assumption* (what condition should be true for the program execution to be well-defined across this edge). In Spec, assertions expressed using the `assuming-do` statement form the UB assumptions. For brevity, we often represent a sequence of instructions with a single edge, e.g., in fig. 10b, the edge `C5`→`C3` represents the path `C5`→`C6`→`C7`→`C8`→`C3`. In such a case, the transfer function of the edge is the composition of the sequence of instructions. Henceforth, We refer to the IR programs as Spec and C directly unless a distinction is necessary.

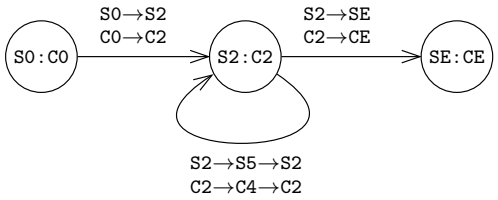
2.3 Equivalence Definition

Given (1) a Spec program specification S, (2) a C implementation C, (3) a precondition Pre that relates the initial inputs $\texttt{Input}_S$ and $\texttt{Input}_C$ to S and C respectively, and (4) a postcondition $Post$ that relates the final outputs $\texttt{Output}_S$ and $\texttt{Output}_C$ of S and C respectively[2]: S and C are *equivalent* if for all possible inputs $\texttt{Input}_S$ and $\texttt{Input}_C$ such that $Pre(\texttt{Input}_S, \texttt{Input}_C)$ holds, S's execution is well-defined on $\texttt{Input}_S$, *and* C's memory allocation requests during its execution on $\texttt{Input}_C$ are successful, then both programs S and C produce outputs such that $Post(\texttt{Output}_S, \texttt{Output}_C)$ holds.

$$Pre(\texttt{Input}_S, \texttt{Input}_C) \wedge (S\ \texttt{def}) \wedge (C\ \texttt{fits}) \Rightarrow Post(\texttt{Output}_S, \texttt{Output}_C)$$

The (S `def`) antecedent states that we are only interested in proving equivalence for well-defined executions of S, i.e., executions that satisfy all assertions expressed using the `assuming-do` statement. Sometimes, the user may be interested in constraining the nature of inputs to C for the purpose of checking equivalence only for *well-defined* inputs. In these cases, we use a combination of Pre and (S `def`) to constrain the execution of C to inputs for which we are interested in proving equivalence. For example, the C library function `strlen(char* str`$_C$`)` is well-defined only if $\texttt{str}_C$ represents a valid null character terminated string. This

[2] $\texttt{Input}_C$ and $\texttt{Output}_C$ include the initial and final memory state of C respectively.



(a) Product-CFG

PC-Pair	Invariants
(S0:C0)	(P) $\mathtt{l}_S \sim \mathtt{Clist}^{\mathtt{lnode}}_{\mathtt{m}}(\mathtt{l}_C)$
(S2:C2)	(I1) $\mathtt{l}_S \sim \mathtt{Clist}^{\mathtt{lnode}}_{\mathtt{m}}(\mathtt{l}_C)$
	(I2) $\mathtt{sum}_S = \mathtt{sum}_C$
(SE:CE)	(E) $\mathtt{ret}_S = \mathtt{ret}_C$

(b) Node Invariants of the Product-CFG

Figure 9: Product-CFG between the CFGs in figs. 8c and 8d. The inductive invariants of the Product-CFG are given in fig. 9b.

includes the assumption that the pointer $\mathtt{str}_C$ may not be null. Since Spec has no notion of pointers, we expose this conditional well-definedness of C strings through an explicit constructor e.g. `SInvalid` for the `String` ADT defined as $\mathtt{String} = \mathtt{SInvalid} \mid \mathtt{SNil} \mid \mathtt{SCons}(\mathtt{i8}, \mathtt{String})$. $(S\ \mathtt{def})$ asserts $\neg(\mathtt{str}_S\ \text{is}\ \mathtt{SInvalid})$ and the precondition Pre contains the relation $(\mathtt{str}_S\ \text{is}\ \mathtt{SInvalid}) \Leftrightarrow (\mathtt{str}_C = 0)$. Hence, $(S\ \mathtt{def})$ and Pre ensures that we compute equivalence only for those executions of S and C where the input strings are well-defined. A similar strategy is employed for other functions as detailed in section 5.2.

The $(C\ \mathtt{fits})$ antecedent states that we prove equivalence under the assumption that C's memory requirements fit within the available system memory i.e., only for those executions of C in which all memory allocation requests (through `malloc` calls) are successful.

The returned values of S and C procedures form their observable outputs. For S, the returned values are explicit and may include ADT values. For C, observables include the returned value along with the implicit memory state at program exit. The postcondition $Post$ relates these outputs of the two programs. In general, the Spec and C sources may contain multiple procedures and these procedures may have calls to each other. In that scenario, we are interested in proving equivalence of each S and C procedure pair.

2.4 Bisimulation Relation

We construct a *bisimulation relation* to identify equivalence between two programs. A bisimulation relation correlates the transitions of S and C in lockstep, such that the lockstep execution ensures identical observable behavior. A bisimulation

relation between two programs can be represented using a *product program* [43] and the CFG representation of a product program is called a *product-CFG*. Figure 9a shows a product-CFG, that encodes the lockstep execution (bisimulation relation) between the CFGs in figs. 8c and 8d.

A node in the product-CFG is formed by pairing nodes of S and C CFGs, e.g., `S2:C2` is formed by pairing `S2` and `C2`. If the lockstep execution of both programs is at node `S2:C2` in the product-CFG, then S's execution is at `S2` and C's execution is at `C2`. The start node `S0:C0` of the product-CFG correlates the start nodes of CFGs of S and C. Similarly, the exit node `SE:CE` correlates the exit nodes of both programs.

An edge in the product-CFG is formed by pairing a *path* (a sequence of edges) in S with a path in C. A product-CFG edge encodes the lockstep execution of its correlated paths. For example, the product-CFG edge (`S2:C2`)→(`S2:C2`) is formed by pairing `S2`→`S5`→`S2` and `C2`→`C4`→`C2` in figs. 8c and 8d, and represents that when S makes the transition `S2`→`S5`→`S2`, C makes the transition `C2`→`C4`→`C2` in lockstep. In general, a product-CFG edge e may correlate a finite path ρ_S in S with a finite path ρ_C in C, written $e = (\rho_S, \rho_C)$. The empty path ϵ in S may be correlated with a finite path in C However, a product-CFG is only well-formed (i.e. represents a valid bisimulation relation) if no loop path in C is correlated with ϵ in S. For example, fig. 11 shows the correlation of ϵ with the paths `C3`→`C4` and `C4`→`C5`. Since the loop path `C3`→`C4`→`C5`→`C3` in C is still correlated with the non-empty path `S3`→`S5`→`S3` in S, it represents a valid bisimulation relation.

At the start node `S0:C0` of the product-CFG in fig. 11, the precondition *Pre* (labeled Ⓟ) ensures equality of input arguments $\mathtt{n}_S$ and $\mathtt{n}_C$ at programs' entry. *Inductive invariants* (labeled Ⓘ) are inferred at each intermediate product-CFG node that relate the values of S with values and memory state of C. The inductive invariants are identified by running an invariant inference algorithm on the product-CFG, which is further discussed in section 4.3. At the exit node `SE:CE` of the product-CFG, the postcondition *Post* (labeled Ⓟ) represents equality of observable outputs and forms our primary proof obligation. Assuming that the precondition *Pre* (Ⓟ) holds at the entry node `S0:C0`, a bisimulation check involves checking that the inductive invariants (Ⓘ) hold too, and consequently the postcondition *Post* (Ⓔ) holds at the exit node `SE:CE`.

2.5 Recursive Relation

In fig. 9b, the precondition (Ⓟ) is an example of a *recursive relation*: "$\texttt{l}_S \sim \texttt{Clist}^{\texttt{lnode}}_{\texttt{m}}(\texttt{l}_C)$" where $\texttt{l}_S$ and $\texttt{l}_C$ represent the input variables to the Spec and C programs respectively, `lnode` is the C `struct` type that contains the `val` and `next` fields, and $\texttt{m}$ is the byte-addressable array representing the current memory state of the C program. $l_1 \sim l_2$ is read *l_1 is recursively equal to l_2* and is semantically equivalent to $l_1 = l_2$. The '$\sim$' simply emphasizes that l_1 and l_2 are (possibly recursive) ADT values. $\texttt{Clist}^{\texttt{lnode}}$ is called a *lifting constructor* that 'lifts' a C pointer value p (pointing to an object of type `struct lnode`) and a C memory state $\texttt{m}$ to a (possibly infinite in case of a circular list) `List` value, and is defined through its *unrolling procedure* as follows:

$$\begin{aligned} U_C : \texttt{Clist}^{\texttt{lnode}}_{\texttt{m}}(p\!:\!\texttt{i32}) = \underline{\texttt{if}}\ p = 0\ \underline{\texttt{then}}\ \texttt{LNil} \\ \underline{\texttt{else}}\ \texttt{LCons}(p \xrightarrow{\texttt{m}}_{\texttt{lnode}} \texttt{val}, \texttt{Clist}^{\texttt{lnode}}_{\texttt{m}}(p \xrightarrow{\texttt{m}}_{\texttt{lnode}} \texttt{next})) \end{aligned} \tag{3}$$

Note the recursive nature of the lifting constructor $\texttt{Clist}^{\texttt{lnode}}$: if the pointer p is zero (i.e. p is a null pointer), then it represents the empty list `LNil`; otherwise it represents the list formed by `LCons`-ing the value stored at $p \xrightarrow{\texttt{m}}_{\texttt{lnode}} \texttt{val}$ in memory $\texttt{m}$ and the list formed by recursively lifting $p \xrightarrow{\texttt{m}}_{\texttt{lnode}} \texttt{next}$ through $\texttt{Clist}^{\texttt{lnode}}$. $\texttt{Clist}^{\texttt{lnode}}_{\texttt{m}}(p)$ allows us to adapt a C linked list (formed by chasing a pointer p in the memory $\texttt{m}$) to a `List` value and compare it with a Spec `List` value for equality.

2.6 Proof Obligations

The counterexample-guided algorithms for construction of the product-CFG and inference of inductive invariants are discussed later in **??**. For now, we discuss the proof obligations that arise from a given product-CFG. Consider the product-CFG in fig. 11. Recall that a bisimulation check involves checking that all inductive invariants and the postcondition *Post* hold at each product-CFG node.

We use relational Hoare triples to express these proof obligations [12, 23]. If ϕ denotes a predicate relating the machine states of S and C, then for a product-CFG

edge $e = (\rho_S, \rho_C)$, $\{\phi_s\}(e)\{\phi_d\}$ denotes the condition: if any machine states σ_S and σ_C of programs S and C are related through precondition $\phi_s(\sigma_S, \sigma_C)$ and the paths ρ_S and ρ_C are executed in S and C respectively, then execution terminates normally in states $\sigma_S^{'}$ (for S) and $\sigma_C^{'}$ (for C) and postcondition $\phi_d(\sigma_S^{'}, \sigma_C^{'})$ holds.

For every product-CFG edge $e = (s \rightarrow d) = (\rho_S, \rho_C)$, we are interested in proving: $\{\phi_s\}(\rho_S, \rho_C)\{\phi_d\}$, where ϕ_s and ϕ_d are the node invariants at the product-CFG nodes s and d respectively. The weakest-precondition transformer is used to translate a Hoare triple $\{\phi_s\}(\rho_S, \rho_C)\{\phi_d\}$ to the following first-order logic formula:

$$(\phi_s \wedge pathcond_{\rho_S} \wedge pathcond_{\rho_C} \wedge ubfree_{\rho_S}) \Rightarrow \mathtt{WP}_{\rho_S,\rho_C}(\phi_d) \tag{4}$$

Here, $pathcond_{\rho_X}$ represent the condition that path ρ is taken in program X and $ubfree_{\rho_S}$ represents the condition that execution of S along path ρ_S is free of undefined behaviour. $\mathtt{WP}_{\rho_S,\rho_C}(\phi_d)$ represents the weakest-precondition of the predicate ϕ_d across the product-CFG edge $e = (\rho_S, \rho_C)$. We will use '`LHS`' and '`RHS`' to refer to the antecedent and consequent of the implication operator '$\Rightarrow$' in eq. (10).

For example, checking that the loop invariant (I2) $\mathtt{l}_S \sim \mathtt{Clist}_{\mathtt{m}}^{\mathtt{lnode}}(\mathtt{l}_C)$ holds at `S2:C2` in fig. 9a requires us to prove the following two proof obligations: (1) $\{\phi_{\mathtt{S0:C0}}\}(\mathtt{S0}{\rightarrow}\mathtt{S2}, \mathtt{C0}{\rightarrow}\mathtt{C2})\{\mathtt{l}_S \sim \mathtt{Clist}_{\mathtt{m}}^{\mathtt{lnode}}(\mathtt{l}_C)\}$ and (2) $\{\phi_{\mathtt{S2:C2}}\}(\mathtt{S2}{\rightarrow}\mathtt{S5}{\rightarrow}\mathtt{S2}, \mathtt{C2}{\rightarrow}\mathtt{C4}{\rightarrow}\mathtt{C2})\{\mathtt{l}_S \sim \mathtt{Clist}_{\mathtt{m}}^{\mathtt{lnode}}(\mathtt{l}_C)\}$. The proof obligation (2) reduces to the following first-order logic proof obligation:

$$\begin{aligned} \mathtt{l}_S \sim \mathtt{Clist}_{\mathtt{m}}^{\mathtt{lnode}}(\mathtt{l}_C) \wedge \mathtt{sum}_S = \mathtt{sum}_C \wedge (\mathtt{l}_S \text{ is } \mathtt{LCons}) \wedge (\mathtt{l}_C \neq 0) \\ \Rightarrow \mathtt{l}_S.\mathtt{next} \sim \mathtt{Clist}_{\mathtt{m}}^{\mathtt{lnode}}(\mathtt{l}_C \xrightarrow[\mathtt{lnode}]{\mathtt{m}} \mathtt{next}) \end{aligned} \tag{5}$$

Due to the presence of recursive relations, these proof queries (e.g., eq. (5)) cannot be solved directly by off-the-shelf solvers and require special handling. The next chapter illustrates our proof discharge algorithm for solving proof queries involving recursive relations.

3 Proof Discharge Algorithm through Illustative Examples

This chapter demonstrates our proof discharge algorithm through examples. We consider proof obligations generated due to invariants shown in table 2 and fig. 9b.

3.1 Properties of Proof Discharge Algorithm

An algorithm that evaluates the truth value of a proof obligation is called a *proof discharge algorithm*. In case a proof discharge algorithm deems a proof obligation to be unprovable, it is expected to return *false* with a set of counterexamples that falsify the proof obligation. A proof discharge algorithm is *precise* if for all proof obligations, the truth value evaluated by the algorithm is identical to the proof obligation's *actual* truth value. A proof discharge algorithm is *sound* if: (a) whenever it evaluates a proof obligation to true, the actual truth value of that proof obligation is also true, and (b) whenever it generates a counterexample, that counterexample must falsify the proof obligation. However, it is possible for a sound proof discharge algorithm to return false (without counterexamples) when the proof obligation was actually provable.

For proof obligations generated by our equivalence checker procedure, it is always safe for a proof discharge algorithm to return false (without counterexamples). Keeping this in mind, our proof discharge algorithm is designed to be *sound*. Conservatively evaluating a proof obligation to false (when it was actually provable) may prevent the equivalence proof from completing successfully. However, importantly, the overall equivalence procedure remains sound i.e. (a) either it successfully finds a valid proof of equivalence (bisimulation relation) or (b) it conservatively returns *unknown*.

Resolving the truth value of a proof obligation that contains a recursive relation such as $\mathtt{l}_S \sim \mathtt{Clist}_{\mathtt{m}}^{\mathtt{lnode}}(\mathtt{l}_C)$ is unclear. Fortunately, the shapes of the proof obligations generated by our equivalence checker are restricted. Our equivalence checking algorithm ensures that, for an invariant $\phi_s = (\phi_{\mathtt{s}}^{\mathtt{1}} \wedge \phi_{\mathtt{s}}^{\mathtt{2}} \wedge ... \wedge \phi_{\mathtt{s}}^{\mathtt{k}})$, at any node s of a product-CFG, if a recursive relation appears in ϕ_s, it must be one of

ϕ_s^1, ϕ_s^2, ..., or ϕ_s^k. We call this the *conjunctive recursive relation* property of an invariant ϕ_s.

A proof obligation $\{\phi_s\}(e)\{\phi_d\}$, where $e = (\rho_S, \rho_C)$, gets lowered using $\texttt{WP}_e(\phi_d)$ (as shown in eq. (10)) to a first-order logic formula of the following form:

$$(\eta_1^l \wedge \eta_2^l \wedge ... \wedge \eta_m^l) \Rightarrow (\eta_1^r \wedge \eta_2^r \wedge ... \wedge \eta_n^r) \tag{6}$$

Thus, due to the conjunctive recursive relation property of ϕ_s and ϕ_d, any recursive relation in eq. (11) must appear as one of η_i^l or η_j^r. To simplify proof obligation discharge, we break a first-order logic proof obligation P of the form in eq. (11) into multiple smaller proof obligations of the form $P_j : (\texttt{LHS} \Rightarrow \eta_{\texttt{j}}^{\texttt{r}})$, for $j = 1..n$. Each proof obligation P_j is then discharged separately. We call this conversion from a bigger query to multiple smaller queries, RHS-*breaking*.

We provide a sound (but imprecise) proof discharge algorithm that converts a proof obligation generated by our equivalence checker into a series of SMT queries. Our algorithm begins by categorizing a proof obligation into one of three types; each type is discussed separately in subsequent sections. The categorization is based on an 'iterative unification and rewriting' procedure, which we describe next. We use an *unroll parameter* k for our categorization.

3.2 Iterative Unification and Rewriting Procedure

We begin with some definitions. An expression e whose top-level constructor is a lifting constructor, e.g., $e = \texttt{Clist}_{\texttt{m}}^{\texttt{lnode}}(\texttt{l}_C)$, is called a *lifted expression*. An expression e of the form $v.\texttt{a}_1.\texttt{a}_2...\texttt{a}_\texttt{n}$ i.e. a variable with *zero* or more *accessor*-operators applied on it, is called a *pseudo-variable*. Note that, a variable v is a pseudo-variable. An expression e in which (a) all accessors (e.g., '_.tail') appear in a pseudo-variable and (b) each *is*-operator (e.g., '_ is LCons') operate on a pseudo-variable, is called a *canonical expression*.

Consider the expression tree of a canonical expression e, formed using the ADT value constructors and the if-then-else sum-deconstruction operator. The leaves of e (also called *atoms* of e) are the pseudo-variables (of scalar and ADT type), the scalar expressions (of Unit, Bool or i<N> types), and lifted expressions.

The *expression path* to a node v in e's tree is the path from the root of e to the node v. The *expression path condition* represents the conjunction of all the `if` conditions (if the `then` branch of taken along the path), or their negation (if the `else` branch is taken along the path). For example, in the expression `if` c `then` a `else` b, the expression path condition of c is `true`, of a is c, and of b is $\neg c$.

When we attempt to unify two expressions, we unify the structures created by the ADT value constructors and the `if`-`then`-`else` operator of their canonical forms. The unification procedure either fails to unify, or it returns tuples (p_1, p_2, a_1, e_2) where atom a_1 at expression path condition p_1 in one expression is correlated with expression e_2 at expression path condition p_2 in the other expression.

For two non-atomic expressions, e_1 and e_2 to unify successfully, it must be true that either the top-level constructor in e_1 and e_2 is the same value constructor (in which case an unification is attempted for each of their children), *or* the top-level constructor in one of e_1 or e_2 is `if`-`then`-`else`.

If the top-level constructor of exactly one of e_1 and e_2 (say e_1) is `if`-`then`-`else`, then e_2 must have a value constructor at its root. In such a case, we *rewrite* e_2 using `if`-`then`-`else` such that the condition of the branch containing e_2 is *true* while all other branches have a *false* condition. For example, we can rewrite `LCons`$(0, l)$ as `if` $false$ `then` $LNil$ `else` $LCons(0, l)$. Next, we unify each child (condition and branch expressions) of the top-level `if`-`then`-`else` operators of (possibly rewritten) e_1 and e_2. Whenever we descend down an `if`-`then`-`else` operator, we keep track of the expression path conditions for both expressions. Recall that the `if`-`then`-`else` operator for an ADT T must have exactly one branch for each value constructor of T. Moreover, the branch associated with the value constructor V must contain an expression whose top-level constructor is V.

If one of e_1 and e_2 (say e_2) is atomic, unification always succeeds and returns (p_2, p_1, e_2, e_1) With each atom of an ADT type, we associate an *unrolling procedure* By definition, an ADT atom is either a pseudo-variable of a lifted expression. Every (pseudo-)variable is associated with its unrolling procedure governed by its ADT. For example, the unrolling procedure for `List` variable l is U_S (eq. (8)). For lifted expressions, the unrolling procedure is given by the its definition, e.g., U_C (eq. (9)) for the lifting constructor $\texttt{Clist}^{\texttt{lnode}}$.

Given two expressions e_a and e_b at expression path conditions p_a and p_b respectively, an *iterative unification and rewriting procedure* $\Theta(e_a, e_b, p_a, p_b)$ is used to identify a set of correlation tuples between the atoms in the two expressions. This iterative procedure begins with an attempt to unify e_a and e_b. If this unification fails, we return a failure for the original expressions e_a and e_b. Else, we obtain correlation tuples between atoms and expressions (with their expression path conditions). If the unification correlates an atom a_1 at expression path condition p_1 with another atom a_2 at expression path condition p_2, we add (p_1, a_1, p_2, a_2) to the final output. Otherwise, if the unification correlates an atom a_1 at expression path condition p_1 to a non-atomic expression e_2 at expression path condition p_2, we *rewrite* a_1 using its unrolling procedure to obtain expression e_1. The unification algorithm then proceeds by unifying e_1 and e_2 through a recursive call to $\Theta(e_1, e_2, p_1, p_2)$. The maximum number of rewrites performed by $\Theta(e_a, e_b, p_a, p_b)$ (before termination) is upper bounded by the sum of number of ADT value constructors in e_a and e_b.

For a recursive relation $l_1 \sim l_2$, we unify l_1 and l_2 through a call to $\Theta(l_1, l_2, true, true)$. If the n tuples obtained after a successful unification are $(p_1^i, a_1^i, p_2^i, a_2^i)$ (for $i = 1 \ldots n$), then the *decomposition* of $l_1 \sim l_2$ is defined as:

$$l_1 \sim l_2 \Leftrightarrow \bigwedge_{i=1}^{n} (p_1^i \wedge p_2^i \rightarrow (a_1^i = a_2^i)) \tag{7}$$

For example, the unification of '$\underline{\texttt{if}}$ c_1 $\underline{\texttt{then}}$ $\texttt{LNil}$ $\underline{\texttt{else}}$ $\texttt{LCons}(0, l_1)$' and '$\underline{\texttt{if}}$ c_2 $\underline{\texttt{then}}$ $\texttt{LNil}$ $\underline{\texttt{else}}$ $\texttt{LCons}(i, \texttt{Clist}_{\mathbb{m}}^{\texttt{lnode}}(l_2))$' yields the correlation tuples: $(true, true, c_1, c_2)$, $(\neg c_1, \neg c_2, 0, i)$ and $(\neg c_1, \neg c_2, l_1, \texttt{Clist}_{\mathbb{m}}^{\texttt{lnode}}(l_2))$. Hence, the recursive relation "$\underline{\texttt{if}}$ c_1 $\underline{\texttt{then}}$ $\texttt{LNil}$ $\underline{\texttt{else}}$ $\texttt{LCons}(0, l_1) \sim$ $\underline{\texttt{if}}$ c_2 $\underline{\texttt{then}}$ $\texttt{LNil}$ $\underline{\texttt{else}}$ $\texttt{LCons}(i, \texttt{Clist}_{\mathbb{m}}^{\texttt{lnode}}(l_2))$" decomposes into $(c_1 = c_2) \wedge (\neg c_1 \wedge \neg c_2 \rightarrow 0 = i) \wedge (\neg c_1 \wedge \neg c_2 \rightarrow l_1 \sim \texttt{Clist}_{\mathbb{m}}^{\texttt{lnode}}(l_2))$. Similarly, the decomposition of $l_1 \sim \texttt{LCons}(42, \texttt{Clist}_{\mathbb{m}}^{\texttt{lnode}}(l_2))$ is given by $(l_1$ is $\texttt{LCons}) \wedge (l_1$ is $\texttt{LCons} \rightarrow l_1.\texttt{val} = 42) \wedge (l_1$ is $\texttt{LCons} \rightarrow l_1.\texttt{next} \sim \texttt{Clist}_{\mathbb{m}}^{\texttt{lnode}}(l_2))$. In case of a failed unification, the *decomposition* is defined to be *false*, e.g., $\texttt{LNil} \sim \texttt{LCons}(0, l)$ decomposes into $false$.

Each conjunctive clause of the form $(p_1^i \wedge p_2^i \rightarrow (a_1^i = a_2^i))$[3] in the decomposition

[3] If a_1^i and a_2^i are ADT values, then we replace $a_1^i = a_2^i$ with $a_1^i \sim a_2^i$.

is called a *decomposition clause*. A decomposition clause may relate only atomic values, i.e., it may relate either (a) two scalars or (b) two ADT variable(s) and/or lifted expression(s). However, we restrict recursive relation invariants to a shape such that each recursive relation in its decomposition strictly relates ADT values to lifted expressions only. This is discussed in more detail along with all other invariant shapes in section 4.3. We *decompose* a recursive relation by replacing it with its decomposition. We *decompose* a proof obligation P to P_D by decomposing all recursive relations in P.

3.3 Categorization of Proof Obligations

We *unroll a recursive relation* $l_1 \sim l_2$ by rewriting the top-level expressions l_1 and l_2 through their unrolling procedures (if possible) and decomposing it. We *unroll an expression* e by unrolling each recursive relation in e. More generally, the k-unrolling of e is found by unrolling the $(k-1)$-unrolling of e recursively. For a decomposed proof obligation P_D : `LHS` $\Rightarrow$ `RHS`, we identify its k-unrolling (say P_K), where k is a fixed parameter called the *unrolling parameter*. After k-unrolling, we *eliminate* those decomposition clauses ($p_1 \wedge p_2 \rightarrow (a_1 = a_2)$) in P_K whose ($p_1 \wedge p_2$) evaluates to false under `LHS` ignoring all recursive relations, yielding an equivalent proof obligation, say P_E. For example, the one-unrolling of P : `LHS` $\Rightarrow l \sim$ $\texttt{Clist}^{\texttt{lnode}}_{\mathbb{m}}(0)$, after elimination, yields P_E : `LHS` $\Rightarrow l$ is `LNil`. We categorize a proof obligation P : `LHS` $\Rightarrow$ `RHS` based on the k-unrolled form of its decomposition (i.e. P_E) as follows:

- Type I: P_E does not contain a recursive relation
- Type II: P_E contains a recursive relation *only* in the `LHS`
- Type III: P_E contains a recursive relation in the `RHS`

The categorization method is *sound* as long as the elimination of decomposition clauses is sound (but possibly not precise). In other words, it is possible that we are unable to eliminate a recursive relation in P_K, due to an imprecise algorithm for elimination of decomposition clauses. However, our proof discharge algorithm remains sound irrespective of such imprecision during categorization. Next, we

describe the algorithm for each of the tree types of proof obligations in sections 3.13 to 3.15.

3.4 Handling Type I Proof Obligations

3.5 Handling Type II Proof Obligations

3.6 Handling Type III Proof Obligations

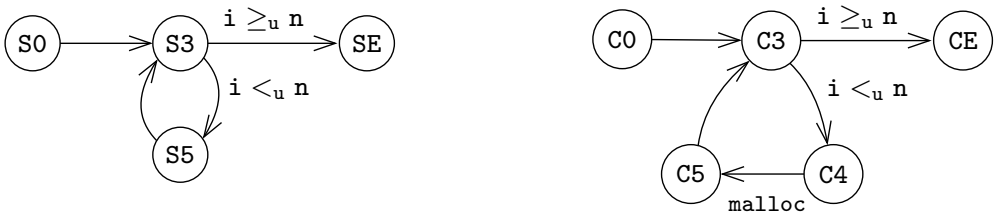
A `List` ADT in the Spec program is defined at line `A0` in fig. 1a. An empty list is represented by the constant `LNil()`[4]; a non-empty list uses the `LCons` constructor to combine its first value (`val:i32`) and the remaining list (`tail:List`). Spec supports `i<N>` (bitvectors of length `N`), `bool`, and `unit` types, also called *scalar types*. Spec's type system prevents the creation of cycles in ADT values. If `l` is an object of type `List`, then to access its constituent values, we may expand (or unroll) `l` to

$$U_S : \texttt{l} \;=\; \underline{\texttt{if}}\ \texttt{l}\ \textit{is}\ \texttt{LNil}\ \underline{\texttt{then}}\ \texttt{LNil}\ \underline{\texttt{else}}\ \texttt{LCons(l.val, l.tail)} \tag{8}$$

In this expanded representation of `l`, the *sum-deconstruction* operator[5] '`if`-`then`-`else`' deconstructs a sum type where the if condition '`l` *is* `Constructor`' checks whether the top level constructor of `l` is '`Constructor`'. If `l` is a non-empty list constructed through `LCons`, then `l.val` and `l.tail` are used to access `l`'s first value and `l`'s tail respectively. The right-hand side of eq. (8) can also be viewed as an executable program that unrolls the input `List` object `l` once and outputs a `List` object constructed from `l`'s constituents — we call eq. (8) the *unrolling procedure* U_S of the `List` ADT. We can similarly define the unrolling procedure for any ADT variable.

[4] `LNil()` represents the application of the nullary constructor `LNil` on the unit value `()`. For brevity, we will simply write `LNil` for `LNil()` henceforth.

[5] The sum-deconstruction operator '`if`-`then`-`else`' for a sum type T must contain exactly one branch for each top-level value constructor of T. For example, '`if`-`then`-`else`' for the `List` type must have exactly two branches of the form `LNil` and `LCons`(e_1,e_2) for some expressions e_1 and e_2.



(a) CFG of Spec Program **(b)** CFG of C Program

Figure 10: CFG representation for Spec and C IRs shown in figs. 2a and 2b

Figures 10a and 10b show the Control-Flow Graph (CFG) representation of the Spec and C programs in figs. 2a and 2b respectively. The CFG nodes represent PC locations of the program, and edges represent transitions through instruction execution. For brevity, we sometimes represent multiple program instructions with a single edge, e.g., in fig. 10b, the edge `C5`→`C3` represents the path `C5`→`C6`→`C7`→`C8`→`C3`. A control-flow edge is associated with an *edge condition* (the condition under which that edge is taken), a *transfer function* (how the program state is mutated if that edge is taken), and a *UB assumption* (what condition should be true for the program execution to be well-defined across this edge). For example, the UB assumption associated with a division instruction in S will encode that the divisor must be non-zero.

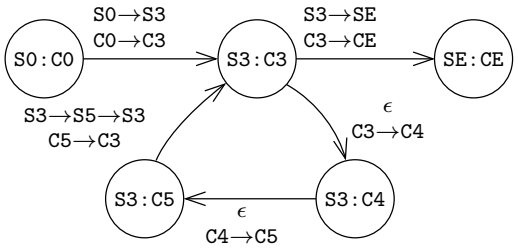


Figure 11: Product-CFG between the CFGs in figs. 10a and 10b

3.7 Product CFG

We construct a *bisimulation relation* to identify equivalence between two programs. A bisimulation relation correlates the transitions of S and C in lockstep, such that this lockstep execution ensures identical observable behavior. An equivalence proof through bisimulation construction can be represented using a *product program* [43] and the CFG of a product program is called a *product-CFG*. Figure 11 shows a

Table 2: Node Invariants for Product-CFG in fig. 11

PC-Pair	Invariants
(S0:C0)	(P) $\mathtt{n}_S = \mathtt{n}_C$
(S3:C3)	(I1) $\mathtt{n}_S = \mathtt{n}_C$ (I2) $\mathtt{i}_S = \mathtt{i}_C$ (I3) $\mathtt{i}_S \leq_u \mathtt{n}_S$ (I4) $\mathtt{l}_S \sim \mathtt{Clist}_{\mathtt{m}}^{\mathtt{lnode}}(\mathtt{l}_C)$
(S3:C4) (S3:C5)	(I5) $\mathtt{n}_S = \mathtt{n}_C$ (I6) $\mathtt{i}_S = \mathtt{i}_C$ (I7) $\mathtt{i}_S <_u \mathtt{n}_S$ (I8) $\mathtt{l}_S \sim \mathtt{Clist}_{\mathtt{m}}^{\mathtt{lnode}}(\mathtt{l}_C)$
(SE:CE)	(E) $\mathtt{ret}_S \sim \mathtt{Clist}_{\mathtt{m}}^{\mathtt{lnode}}(\mathtt{ret}_C)$

product-CFG, that encodes the lockstep execution (bisimulation relation) between the CFGs in figs. 10a and 10b.

A node in the product-CFG is formed by pairing nodes of S and C CFGs, e.g., `(S3:C5)` is formed by pairing `S3` and `C5`. If the lockstep execution is at node `(S3:C5)` in the product-CFG, then S's execution is at `S3` and C's execution is at `C5`. The start node `(S0:C0)` of the product-CFG correlates the start nodes of the CFGs of both programs. Similarly, the exit node `(SE:CE)` correlates the exit nodes of the CFGs of both programs.

An edge in the product-CFG is formed by pairing a path (a sequence of edges) in S with a path in C. A product-CFG edge encodes the lockstep execution of correlated transitions (or paths). For example, the product-CFG edge `(S3:C5)`→`(S3:C3)` is formed by pairing the `(S3`→`S5`→`S3)` and `C5`→`C3` in figs. 10a and 10b, and represents that when S makes a transition `(S3`→`S5`→`S3)`, then C makes the transition `C5`→`C3` in lockstep. The edge `(S3:C3)`→`(S3:C4)` correlates the ϵ path (no transition) in S with `C3`→`C4` in C. In general, a product-CFG edge e may correlate a finite path ρ_S in S with a finite path ρ_C in C, written $e = (\rho_S, \rho_C)$.

At the start node `(S0:C0)` of the product-CFG, the precondition Pre (labeled (P)) ensures the equality of input arguments $\mathtt{n_S}$ and $\mathtt{n_C}$ at programs' entry. Inductive invariants are inferred at each product-CFG node that relate the variables of S with variables and memory locations of C. The inductive invariants are identified by running an invariant inference algorithm on the product-CFG, which is further discussed in section 4.3. The inductive invariants for our example are shown in table 2. For example, at node `(S3:C5)` in fig. 11, $\mathtt{i}_S = \mathtt{i}_C$ is an inductive invariant. If the inferred invariants ensure that the postcondition $Post$ holds at the exit node `(SE:CE)` (labeled (E)), we have shown equivalence of both programs.

3.8 Recursive relations

TODO:try to update the intro to recursive relations(first line)

In table 2, the relation between programs' variables at product-CFG nodes `S3:C3`, `S3:C4` and `S3:C5` is encoded as a recursive relation: "$\texttt{l}_S \sim \texttt{Clist}_m^{\texttt{lnode}}(\texttt{l}_C)$" where $\texttt{l}_S$ and $\texttt{l}_C$ represent the `l` variables in the Spec and C programs respectively, `lnode` represents the C `struct` type that contains the `val` and `next` fields, and m represents a byte-addressable array representing the current memory state of the C program. $l_1 \sim l_2$ is read l_1 *is recursively equal to* l_2, i.e., l_1 and l_2 are isomorphic and have equal values[6]. The *lifting constructor* $\texttt{Clist}_m^{\texttt{lnode}}(\texttt{p})$ is a constructor that *lifts* the C pointer value `p` (pointing to an object of `struct lnode`) and the C memory state m to a Spec `List` value. $\texttt{Clist}_m^{\texttt{lnode}}(\texttt{p})$ is defined through its unrolling procedure as:

$$\begin{aligned} U_C : \texttt{Clist}_m^{\texttt{lnode}}(\texttt{p} : \texttt{i32}) = \underline{\texttt{if}}\ (\texttt{p} == \texttt{0})\ \underline{\texttt{then}}\ \texttt{LNil} \\ \underline{\texttt{else}}\ \texttt{LCons}(p \xrightarrow{m}_{\texttt{lnode}} \texttt{val},\ \texttt{Clist}_m^{\texttt{lnode}}(p \xrightarrow{m}_{\texttt{lnode}} \texttt{next})) \end{aligned} \tag{9}$$

By construction, this unrolling procedure U_C is isomorphic to `List`'s unrolling procedure U_S in eq. (8). "$p \xrightarrow{m}_{\texttt{s}} \texttt{f}$" represents the field '`f`' of the 'struct `s`' object pointed-to by pointer '`p`' in memory state 'm'. When represented in C-like syntax, '$p \xrightarrow{m}_{\texttt{s}} \texttt{f}$' is equivalent to "`*((typeof s.f)*)(&m[p+offsetof(s,f)])`", i.e., the expression '$p \xrightarrow{m}_{\texttt{s}} \texttt{f}$' returns the bytes in the memory array 'm' starting at address '`p+offsetof(s,f)`' and interpreted as an object of type '`typeof s.f`'.

Note the recursive nature of the lifting value constructor `Clist`: if the pointer `p` of type `i32`[7] is zero (i.e. `p` is a null pointer), then this represents the empty list (`LNil`); otherwise it represents the list formed by `LCons`-ing the value stored at `p->val` in memory m and the list formed by recursively lifting `p->next` using `Clist` in memory m. The recursive lifting constructor `Clist` allows us to compare C values and Spec values for equality. In general, an equality relation between two (possibly recursive) ADT values is called a *recursive relation*. However in the context of bisimulation, we will only consider recursive relations between Spec values (such as variables) and lifted C values (lifted using a lifting constructor such

[6] $l_1 \sim l_2$ and $l_1 = l_2$ are equivalent — the former emphasizes the recursive nature of the values being compared.

[7] The IR lowers integers and pointers in C to bitvectors of type `i<N>`. e.g., `i32` is a 32-bit bitvector type.

as `Clist`).

We later discuss in section 4.2 how a product-CFG can be constructed automatically through a counterexample-guided search. Before that, we discuss the proof obligations that arise from a given product-CFG. Consider the product-CFG in fig. 11. Assuming that the precondition Ⓟ holds at the entry node `S0:C0` of this product-CFG, a bisimulation check involves checking that the invariants at the other product-CFG nodes hold too, and consequently the postcondition Ⓔ holds at the exit node `SE:CE`. Recall that the precondition Ⓟ and the postcondition Ⓔ are provided by the user, but all the other invariants are inferred automatically.

3.9 Proof Obligations

We use relational Hoare triples to express these proof obligations [12, 23]. If ϕ denotes a predicate relating the machine states of programs S and C, then for a product-CFG edge $e = (\rho_S, \rho_C)$, $\{\phi_s\}(e)\{\phi_d\}$ denotes the condition: if the machine states σ_S and σ_C of programs S and C are related through precondition $\phi_s(\sigma_S, \sigma_C)$ and paths ρ_S and ρ_C are executed in S and C respectively (implying the path conditions hold), then execution terminates normally in states $\sigma_S^{'}$ (for S) and $\sigma_C^{'}$ (for C) where postcondition $\phi_d(\sigma_S^{'}, \sigma_C^{'})$ hold. $\{\phi_s\}(e)\{\phi_d\}$ can also be written as $\{\phi_s\}(\rho_S, \rho_C)\{\phi_d\}$.

For every product-CFG edge $e = (s \rightarrow d) = (\rho_S, \rho_C)$ in fig. 11, we thus need to prove $\{\phi_s\}(\rho_S, \rho_C)\{\phi_d\}$, where ϕ_s and ϕ_d are the node invariants (shown in table 2) at nodes s and d of the product-CFG respectively. The weakest-precondition transformer is used to translate a Hoare triple $\{\phi_s\}(\rho_S, \rho_C)\{\phi_d\}$ to the following first-order logic formula:

$$(\phi_s \wedge pathcond_{\rho_S} \wedge pathcond_{\rho_C} \wedge ubfree_{\rho_S}) \Rightarrow \mathtt{WP}_{\rho_S,\rho_C}(\phi_d) \tag{10}$$

Here, $pathcond_{\rho_X}$ represents the condition that path ρ_X is taken in program X. $ubfree_{\rho_S}$ represents the condition that the execution of program S along path ρ_S is free of undefined behavior. $\mathtt{WP}_{\rho_S,\rho_C}(\phi_d)$ represents the weakest-precondition of the predicate ϕ_d across the product-CFG edge $e = (\rho_S, \rho_C)$. We will use "`LHS`" and "`RHS`" to refer to the left and right hand sides of the implication operator "$\Rightarrow$" in eq. (10).

3.10 Proof Discharge Algorithm and Its Soundness

We call an algorithm that evaluates the truth value of a proof obligation, a *proof discharge algorithm*. In case a proof discharge algorithm deems a proof obligation to be unprovable, it is expected to return *false* with a set of counterexamples that falsifies the proof obligation. A proof discharge algorithm is *precise* if for all proof obligations, the truth value evaluated by the algorithm is identical to the proof obligation's *actual* truth value. A proof discharge algorithm is *sound* if: (a) whenever it evaluates a proof obligation to true, the actual truth value of that proof obligation is also true, and (b) whenever it generates a counterexample, that counterexample must falsify the proof obligation. However, it is possible for a sound proof discharge algorithm to return false (without a counterexample) when the proof obligation was actually true.

For the proof obligations generated by our equivalence procedure, it is always safe for a proof discharge algorithm to return false (without a counterexample). If a proof discharge algorithm conservatively evaluates a proof obligation to false (when it was actually true), it may prevent the overall equivalence proof from completing successfully; however, importantly, the overall equivalence procedure remains sound.

Resolving the truth value of a proof obligation that contains a recursive relation such as $\mathtt{l_s} \sim \mathtt{Clist}_{\mathtt{m}}^{\mathtt{lnode}}(\mathtt{l_c})$ is unclear. Fortunately, the shapes of the proof obligations generated by our equivalence checking algorithm are restricted, which makes it possible to soundly resolve these proof obligations.

Our equivalence checking algorithm ensures that, for an invariant $\phi_s = (\phi_{\mathtt{s}}^{\mathtt{1}} \wedge \phi_{\mathtt{s}}^{\mathtt{2}} \wedge ... \wedge \phi_{\mathtt{s}}^{\mathtt{k}})$, at any node s of a product-CFG, if a recursive relation appears in ϕ_s, it must be one of ϕ_s^1, ϕ_s^2, ..., or ϕ_s^k. We call this the *conjunctive recursive relation* property of an invariant ϕ_s.

A proof obligation $\{\phi_s\}(e)\{\phi_d\}$, where $e = (\rho_S, \rho_C)$, gets lowered using $\mathtt{WP}_e(\phi_d)$ (as shown in eq. (10)) to a first-order logic formula of the following form:

$$(\eta_{\mathtt{1}}^{\mathtt{l}} \wedge \eta_{\mathtt{2}}^{\mathtt{l}} \wedge ... \wedge \eta_{\mathtt{m}}^{\mathtt{l}}) \Rightarrow (\eta_{\mathtt{1}}^{\mathtt{r}} \wedge \eta_{\mathtt{2}}^{\mathtt{r}} \wedge ... \wedge \eta_{\mathtt{n}}^{\mathtt{r}}) \tag{11}$$

In this formula, the `LHS` and `RHS` are written as conjunctions of η_i^l and η_j^r respectively (for $1 \leq i \leq m$, $1 \leq j \leq n$). Each η_j^r relation is obtained from $\mathtt{WP}_e(\phi_d^j)$,

where $\phi_d = (\phi_{\mathtt{d}}^{\mathtt{1}} \wedge \phi_{\mathtt{d}}^{\mathtt{2}} \wedge ... \wedge \phi_{\mathtt{d}}^{\mathtt{n}})$. Thus, due to the conjunctive recursive relation property of ϕ_s and ϕ_d, any recursive relation in eq. (11) must appear as one of η_i^l or η_j^r.

To simplify proof obligation discharge, we break a first-order logic proof obligation P of the form in eq. (11) into multiple smaller proof obligations of the form $P_j : (\mathtt{LHS} \Rightarrow \eta_{\mathtt{j}}^{\mathtt{r}})$, for $j = 1..n$. Each proof obligation P_j is then discharged separately. We call this conversion from a bigger query to multiple smaller queries, *RHS-breaking*.

3.11 Iterative Unification and Unrolling

We begin with some definitions. An expression e whose top-level constructor is a lifting constructor, e.g., $e = \mathtt{Clist}_{\mathtt{m}}^{\mathtt{lnode}}(\mathtt{l_C})$, is called a *lifted expression*. An expression e of the form $\mathtt{v.a}_1\mathtt{.a}_2\mathtt{...a}_n$ i.e. a variable with *zero* or more *product deconstruction* operators applied on it, is called a *pseudo-variable*. By definition, variables are pseudo-variables. An expression e in which (1) all product deconstructors (e.g. '`.tail`') appear as part of a *pseudo-variable* and (2) each *sum-is* operator (e.g. '*is* `LCons`') operate on a *pseudo-variable*, is called a *canonical expression*.

Consider the expression tree of a canonical expression e of ADT T, formed using the ADT value constructors and the `if-then-else` sum-deconstruction operator. The leaves of e (also called atoms of e) are the pseudo-variables (of scalar or ADT type), the scalar expressions (of `unit`, `bool`, or `i<N>` types), and lifted expressions.

The *expression path* to a node v in e's tree is the path from the root of e to that node v. The *expression path condition* represents the conjunction of all the if conditions (if the then branch is taken on the expression path), or their negation (if the else branch is taken on the expression path) seen on the expression path. For example, in expression `if (c) then a else b`, the expression path condition of `c` is `true`, of `a` is `c`, and of `b` is $\neg$`c`.

When we attempt to unify two expressions, we unify the structures created by the value constructors and the 'if-then-else' operator of their canonical forms. The unification procedure either fails to unify, or it returns tuples (p_1, p_2, a_1, e_2) where atom a_1 at expression path condition p_1 in one expression is correlated with expression e_2 at expression path condition p_2 in the other expression.

For two non-atomic expressions e_1 and e_2 to unify successfully, it must be true that either the top-level node in both e_1 and e_2 have the same value constructor (in which case a unification is attempted for each of the children of the top-level constructor), *or* the top-level node in one of e_1 or e_2 is if-then-else. If the top-level node of e_1 is "if (c) then $e_1^{\texttt{then}}$ else $e_1^{\texttt{else}}$, we attempt to unify both $e_1^{\texttt{then}}$ and $e_1^{\texttt{else}}$ with e_2 and return success iff any of these attempts succeed (similarly for e_2). Whenever we descend down an if-then-else operator, we conjunct the corresponding if condition (for $e_1^{\texttt{then}}$) or its negation (for $e_1^{\texttt{else}}$) to the respective expression path condition. If one of e_1 and e_2 (say e_2) is atomic, unification always succeeds and returns (p_2, p_1, e_2, e_1).

With each atom of an ADT type, we associate an unrolling procedure. By definition, an ADT atom is either a pseudo-variable or a lifted expression. Every (pseudo-)variable is associated with its unrolling procedure as governed by its ADT. For example, the unrolling procedure for a Spec variable `l` of `List` type is U_S (eq. (8)). For lifted expressions, the unrolling procedure is given by the definition of the lifting constructor such as U_C (eq. (9)) for the lifting constructor `Clist`.

Given two expressions e_a and e_b of an ADT T at expression path conditions p_a and p_b respectively, an *iterative unrolling and unification procedure* $\Theta(e_a, e_b, p_a, p_b)$ is used to identify a set of correlation tuples between the atoms in the two expressions. This iterative procedure proceeds by attempting to unify e_a and e_b. If this unification fails, we return a unification failure for the original expressions e_a and e_b. Else, we obtain correlation between atoms and expressions (with their expression path conditions). If the unification correlates an atom a_1 at expression path condition p_1 with another atom a_2 at expression path condition p_2, we add (p_1, a_1, p_2, a_2) to the final output. If the unification correlates an atom a_1 at expression path condition p_1 to a non-atomic expression e_2 at expression path condition p_2, we unroll a_1 once using its unrolling procedure to obtain expression e_1. The unification algorithm then proceeds by unifying e_1 and e_2 through a recursive call to $\Theta(e_1, e_2, p_a \wedge p_1, p_b \wedge p_2)$. The maximum number of unrollings performed by $\Theta(e_a, e_b, p_a, p_b)$ (before converging) is upper bounded by the sum of number of ADT value constructors in e_a and e_b.

If a proof obligation involves a recursive relation $e_a \sim e_b$, we unify

e_a and e_b through a call to $\Theta(e_a, e_b, \texttt{true}, \texttt{true})$. For example, the unification of "`if (c1) then LNil else LCons(0, lS)`" and "`if (c2) then LNil else LCons(0, Clist`$^{\texttt{lnode}}_m$`(lC))`" yields the correlation tuples: `(c1,(),c2,())`, `(¬c1,0,¬c2,0)` and `(¬c1,lS,¬c2,`$Clist^{\texttt{lnode}}_m$`(lC))`.

If the set of n tuples obtained after a successful unification of $e_a \sim e_b$ are $(p_1^i, a_1^i, p_2^i, a_2^i)$ (for $i = 1 \ldots n$), then $e_a \sim e_b \Leftrightarrow \bigwedge_{i=1}^{n}((p_1^i = p_2^i) \wedge ((p_1^i \wedge p_2^i) \rightarrow (a_1^i = a_2^i)))$[8]. We call $\bigwedge_{i=1}^{n}((p_1^i = p_2^i) \wedge ((p_1^i \wedge p_2^i) \rightarrow (a_1^i = a_2^i)))$ the *decomposition* of $e_a \sim e_b$. Each conjunctive clause of one of the forms $(p_1^i = p_2^i)$ and $((p_1^i \wedge p_2^i) \rightarrow (a_1^i = a_2^i))$ in this decomposition is called a *decomposition clause*. A decomposition clause may relate only atomic values, i.e., in the decomposed form, all recursive relations relate only ADT variables and/or lifted expressions. The decomposition for a failed unification is defined to be `false`. We *decompose* a recursive relation by replacing it with its decomposition. We *decompose* a proof obligation P by decomposing all recursive relations in P.

3.12 k-unrolling with respect to an unrolling procedure

We can *unroll an expression e with respect to an unrolling procedure U* by substituting all occurrences of `LHS` in U by its unrolled version (`RHS` in U) and decomposing it. An expression e is unrolled (without specifying an unrolling procedure U) by unrolling it with respect to the unrolling procedures associated with each of its ADT atoms. A k-unrolling of an expression e is obtained by unrolling a $(k-1)$ unrolling of e.

For a first-order logic proof obligation P : `LHS` $\Rightarrow$ `RHS`, we identify a k-unrolling of P (for a fixed unrolling parameter k). After unrolling, we eliminate those decomposition clauses $(p_1^i \wedge p_2^i) \rightarrow (a_1^i = a_2^i)$ whose path condition $(p_1^i \wedge p_2^i)$ evaluates to `false` under the `LHS` ignoring all recursive relations. For example, the one-unrolling of P : `LHS` $\Rightarrow$ `ls` $\sim$ `Clist`$^{\texttt{lnode}}_{\texttt{m}}$`(0)` after simplification yields P' : `LHS` $\Rightarrow$ `ls is LNil`. Note that unlike P, P' does not contain a recursive relation in its `RHS`. We categorize a proof obligation P based on this k-unrolled form of P as follows:[9]:

[8] If a_1^i and a_2^i are ADT values, then we replace $a_1^i = a_2^i$ with $a_1^i \sim a_2^i$.

[9] TODO: make this a para, not a footnote. In general, checking whether a path condition

- Type I: k-unrolling of P does not contain recursive relations
- Type II: k-unrolling of P contains recursive relations only in the `LHS`
- Type III: k-unrolling P contains recursive relations in the `RHS`

3.13 Type I: k-unrolling of P does not contain recursive relations

In fig. 11, consider a proof obligation generated across the product-CFG edge $(\texttt{S0} : \texttt{C0}) \rightarrow (\texttt{S3} : \texttt{C3})$ while checking if the (I4) invariant, $l_S \sim Clist_m^{\texttt{lnode}}(l_C)$, holds at `(S3:C3)`: $\{\phi_{\texttt{S0:C0}}\}(\texttt{S0} \rightarrow \texttt{S3}, \texttt{C0} \rightarrow \texttt{C3})\{l_S \sim Clist_m^{\texttt{lnode}}(l_C)\}$. The precondition $\phi_{\texttt{S0:C0}} = (n_S = n_C)$ does not contain a recursive relation. When lowered to first-order logic through $\texttt{WP}_{\texttt{S0}\rightarrow\texttt{S3},\texttt{C0}\rightarrow\texttt{C3}}$, this translates to "$(n_S = n_C) \Rightarrow (\texttt{LNil} \sim \texttt{Clist}_{\texttt{m}}^{\texttt{lnode}}(\texttt{0}))$". Here, `LNil` is obtained for $\texttt{l}_S$ and `0` (null) is obtained for $\texttt{l}_C$. The one-unrolled form of this proof obligation yields $(n_S = n_C) \Rightarrow \texttt{LNil} \sim \texttt{LNil}$ which trivially resolves to true.

Consider another example of a proof obligation, $\{\phi_{\texttt{S0:C0}}\}(\texttt{S0} \rightarrow \texttt{S3} \rightarrow \texttt{S5} \rightarrow \texttt{S3}, \texttt{C0} \rightarrow \texttt{C3})\{l_S \sim Clist_m^{\texttt{lnode}}(l_C)\}$. Notice, we have changed the path in S to $\texttt{S0} \rightarrow \texttt{S3} \rightarrow \texttt{S5} \rightarrow \texttt{S3}$ here. In this case, the corresponding first-order logic condition evaluates to: "$(n_S = n_C) \Rightarrow (\texttt{LCons}(\texttt{0}, \texttt{LNil}) \sim \texttt{Clist}_{\texttt{m}}^{\texttt{lnode}}(\texttt{0}))$". One-unrolling of this proof obligation converts $\texttt{Clist}_{\texttt{m}}^{\texttt{lnode}}(\texttt{0})$ to `LNil`, and decomposes `RHS` into `false`. The proof obligation is further discharged using an SMT solver which provides a counterexample (model) that evaluates the formula to false. For example, the counterexample $\texttt{\{ n}_S \mapsto \texttt{42, n}_C \mapsto \texttt{42 \}}$ evaluates this formula to false. These counterexamples assist in faster convergence of our invariant inference and correlation search procedures (as we will discuss later in sections 4.2 and 4.3).

Thus, we unify the structure and values of the Spec objects on both sides of the $\sim$ operator (after k unrollings), and discharge the resulting proof obligations (that

is `false` requires a proof discharge for each clause. Instead, we use a syntactic simplifier to identify these decomposition clauses. A more sophisticated approach would likely improve the completeness of the overall algorithm at the cost of performance. Nevertheless, this imprecise categorization is sound. Similarly, a higher value for the parameter k would increase completeness of categorization and the overall proof discharge algorithm.

```
A0: type List = LNil | LCons (val:i32, tail:List).
A1:
A2: fn sum_list_impl (l:List) (sum:i32) : i32 =
A3:    match l with
A4:    | LNil => sum
A5:    | LCons(x, rest) => sum_list_impl(rest, sum + x).
A6:
A7: fn sum_list (l:List) : i32 = sum_list_impl(l, 0_i32).
```

(a) Spec Program

```
B0:  typedef struct lnode {
B1:    unsigned val; struct lnode* next; } lnode;
B2:
B3:  unsigned sum_list(lnode* l) {
B4:    unsigned sum = 0;
B5:    while (l) {
B6:      sum += l→val;
B7:      l = l→next;
B8:    }
B9:    return sum;
B10: }
```

(b) C Program

Figure 12: Spec and C Programs traversing a Linked List.

relate bitvector and array values) using an SMT solver. Please refer to `Chapter XXX` of the thesis for the intricacies of (a) translation of the formlua to SMT logic and (b) reconstruction of counterexamples from the models returned by the SMT solver. Assuming a capable enough SMT solver, all proof obligations in Type I can be discharged precisely, i.e., we can always decide whether P evaluates to true or false. If it evaluates to false, we also obtain a counterexample.

3.14 Type II: k-unrolling of P contains recursive relations only in the LHS

Consider the pair of programs in figs. 12a and 12b that traverse a list to compute the sum of all elements. The corresponding product-CFG and its node invariants that ensure observable equivalence are shown in figs. 9a and 9b.

Consider the proof obligation originating due to (I2) invariant across edge $(\texttt{S2:C2}) \rightarrow (\texttt{S2:C2})$ in fig. 12: $\{\phi_{\texttt{S2:C2}}\}(\texttt{S2} \rightarrow \texttt{S5} \rightarrow \texttt{S2}, \texttt{C2} \rightarrow \texttt{C4} \rightarrow \texttt{C2})\{sum_S = sum_C\}$, where the node invariant $\phi_{\texttt{S2:C2}}$ contains the recursive relation $l_S \sim$

$Clist_m^{\texttt{lnode}}(l_C)$. The corresponding (simplified) first-order logic condition for this proof obligation is: $(l_S \sim Clist_m^{\texttt{lnode}}(l_C) \wedge \texttt{sum}_S = \texttt{sum}_C \wedge \neg(\texttt{l}_\texttt{S}\ is\ \texttt{LNil}) \wedge \texttt{l}_\texttt{C} \neq \texttt{0}) \Rightarrow ((\texttt{sum}_S + l_S.val) = (sum_C + l \xrightarrow{m}_{\texttt{lnode}} \texttt{val}))$. We fail to remove the recursive relation on the LHS even after k-unrolling for any finite depth k because both sides of $\sim$ represents list values of arbitrary length. In such a scenario, we do not know of an efficient SMT encoding for the recursive relation $(l_S \sim Clist_m^{\texttt{lnode}}(l_C))$. Ignoring this recursive relation will incorrectly (although soundly) evaluate the proof obligation to false; however, for a successful equivalence proof, we need the proof discharge algorithm to evaluate it to true. Let's call this requirement (R1).

Now, consider the proof obligation formed by correlating two iterations of the loop in program S with one iteration of the loop in program C, $\{\phi_{\texttt{S2:C2}}\}(\texttt{S2} \rightarrow \texttt{S5} \rightarrow \texttt{S2} \rightarrow \texttt{S5} \rightarrow \texttt{S2}, \texttt{C2} \rightarrow \texttt{C4} \rightarrow \texttt{C2})\{sum_S = sum_C\}$. Similar to the last proof obligation, its equivalent first-order logic condition contains a recursive relation in the LHS. Clearly, this proof obligation is false. Whenever a proof obligation evaluates to false, we expect an ideal proof discharge algorithm to generate a counterexample that falsifies the proof obligation. Let's call this requirement (R2). Recall that such counterexamples help in faster convergence of our invariant inference and correlation algorithms.

To tackle requirements (R1) and (R2), our proof discharge algorithm converts the original proof obligation $P : \{\phi_s\}(e)\{\phi_d\}$ into two approximated proof obligations: $(P_{\texttt{pre-o}} : \{\phi_s^{o_{d_1}}\}(e)\{\phi_d\})$ and $(P_{\texttt{pre-u}} : \{\phi_s^{u_{d_2}}\}(e)\{\phi_d\})$. Here $\phi_s^{o_{d_1}}$ and $\phi_s^{u_{d_2}}$ represent the over- and under-approximated versions of precondition ϕ_s respectively, and d_1 and d_2 represent *depth* parameters that indicate the degree of over- and under-approximation. To explain our over- and under-approximation scheme, we first introduce the notion of the *depth of an ADT value*.

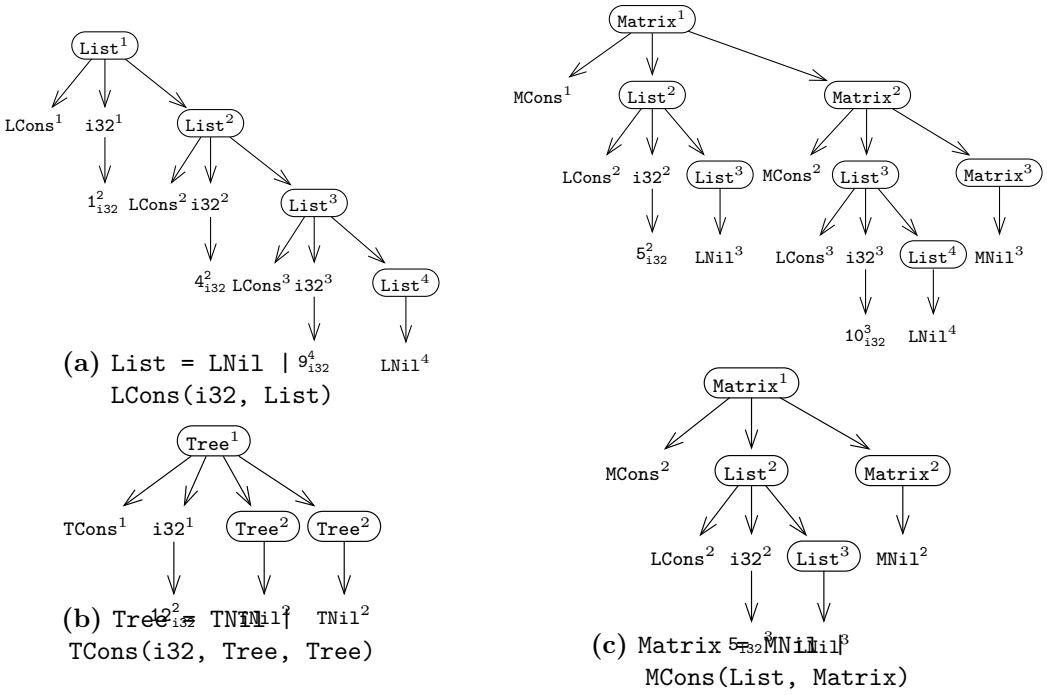


Figure 13: Parsed Trees with depths (shown as superscript) for three values, each of ADT type `List`, `Tree`, and `Matrix` respectively.

3.14.1 Depth of an ADT value

To define the depth of an ADT value, we view the ADT as a context-free grammar:

- The set of terminals consist of (a) scalar values, e.g., `42` for the `i32` type, and (b) the ADT value constructors (e.g., `LNil`, `LCons`).
- The set of non-terminals are all the scalar type identifiers (`i32`, `bool`, `unit`) and all the ADT type identifiers (e.g., `List`, `Tree`, `Matrix2D`).
- Productions of this context-free grammar specify the values that may be constructed for each type identifier: (a) For each basic type (e.g. `i32`), we add productions from its type identifier to all values of that type, e.g., `i32` $\rightarrow$ `0|1|2|...|`$(2^{32}-1)$; (b) For an ADT, each constructor represents a separate production rule, e.g., ADT `List` has two productions: `List` $\rightarrow$ `LNil` | `LCons i32 List`.
- The start non-terminal is the top-level ADT identifier (e.g., `List`).

In this context-free grammar interpretation of an ADT, a value of this ADT type can be viewed as a *parse tree* (also called a derivation tree) of the grammar. The *depth* of a node in this parse tree is the number of ADT identifiers (but not scalar type identifiers) in the path from the root node to the node representing that terminal value (both inclusive). Figure 13 shows examples of values of the `List`, `Tree`, and `Matrix` ADTs, and the depth values of the parse tree nodes. The *depth of an ADT value* is the maximum depth of any node in the parse tree of that value.

3.14.2 Overapproximation and Underapproximation

To overapproximate (underapproximate) a precondition ϕ, each conjunctive recursive relation in ϕ is overapproximated (underapproximated) individually.

The d-depth overapproximated version of a recursive relation $l_1 \sim l_2$ is written as $l_1 \sim_d l_2$, where $\sim_d$ represents the condition that the two ADT values l_1 and l_2 are *recursively equal up to depth d.* i.e., all *terminals* at depth $\leq d$ in the parse trees of both values are identical; however, terminals at depths $> d$ can have different values. $l_1 \sim_d l_2$ (for finite d) is a weaker condition than $l_1 \sim l_2$ (overapproximation); $l_1 \sim l_2$ is equivalent to $l_1 \sim_\infty l_2$.

The d-depth underapproximated version of a recursive relation $l_1 \sim l_2$ is written as $l_1 \approx_d l_2$, where $\approx_d$ represents the condition that the two ADT values l_1 and l_2 are *recursively equal and bounded to depth d*, i.e., l_1, l_2 have a maximum depth $\leq d$ *and* they are recursively equal up to depth d. Thus, $l_1 \approx_d l_2$ is equivalent to $(l_1 \sim_d l_2) \wedge \Gamma_d(l_1) \wedge \Gamma_d(l_2)$, where $\Gamma_d(l)$ represents the condition that the maximum depth of l is d. $l_1 \approx_d l_2$ (for finite d) is a stronger condition than $l_1 \sim l_2$ (underapproximation) as it ensures both equality and max-depth of both values. For arbitrary depths a and b ($a \leq b$), the approximate versions of recursive relation are related as follows:

$$l_1 \approx_a l_2 \Rightarrow l_1 \approx_b l_2 \Rightarrow l_1 \sim l_2 \Rightarrow l_1 \sim_b l_2 \Rightarrow l_1 \sim_a l_2$$

3.14.3 SMT encoding of overapproximate and underapproximate proof obligations

Unlike the original recursive relation $l_1 \sim l_2$, $l_1 \sim_d l_2$ and $l_1 \approx_d l_2$ can be encoded using SMT as shown below:

- $l_1 \sim_d l_2$ is equivalent to the condition that the parse tree structures of the two values l_1 and l_2 (after d-unrolling) are isomorphic till depth d *and* the corresponding values in both (d-depth) isomorphic structures are also equal. $l_1 \sim_d l_2$ can be identified through a d-depth bounded version of iterative unrolling and unification procedure described in section 3.11 followed by keeping only those correlation tuples that equate scalar expressions during decomposition. Along with the expressions and path conditions, we also keep track of the depth parameter which is incremented by one for each successfully unified ADT value constructors. We terminate early if the current depth is greater than d. In case both expressions are atoms, we eagerly unroll both sides and unify further as long as the current depth is within bounds.

  For example, the condition $l \sim_1 \texttt{Clist}_m^{\texttt{lnode}}(\texttt{p})$ can be computed through one-depth bounded iterative unrolling and unification. During decomposition, keeping only correlation tuples that equate scalar expressions, the condition above reduces to the SMT-encodable predicate:

$$((l \ \ is \ \ \texttt{LNil}) \Leftrightarrow (\texttt{p} == 0)) \wedge (\neg(l \ \ is \ \ \texttt{LNil}) \Rightarrow (l.\texttt{val} = p \xrightarrow{m}_{\texttt{lnode}} \texttt{val}))$$

- $\Gamma_d(l)$ is equivalent to the condition that the parse-tree nodes at depths $> d$ are unreachable. This is achieved by unrolling a recursive relation till depth d and then asserting the unreachability of if-then-else paths that reach nodes with depth $> d$ (by checking the satisfiability of their expression path conditions). For example, for a `List` value l, the condition $\Gamma_2(l)$ is equivalent to $(l \ \ is \ \ \texttt{LNil}) \vee (\neg(l \ \ is \ \ \texttt{LNil}) \wedge (l.\texttt{tail} \ \ is \ \ \texttt{LNil}))$. Similarly, $\Gamma_2(\texttt{Clist}_m^{\texttt{lnode}}(\texttt{p}))$ is equivalent to $(\texttt{p} = 0) \vee (\neg(\texttt{p} \neq 0) \wedge (\texttt{p} \xrightarrow{m}_{\texttt{lnode}} \texttt{next} = 0))$.

3.14.4 Proof discharge algorithm for Type II proof obligations

Thus, for a *Type II* proof obligation $P : \{\phi_s\}(e)\{\phi_d\}$, we first submit the proof obligation $(P_{\texttt{pre-o}} : \{\phi_s^{o_{d_1}}\}(e)\{\phi_d\})$ to the SMT solver. Recall that the precondition $\phi_s^{o_{d_1}}$ is the overapproximated version of ϕ_s. If the SMT solver evaluates $P_{\texttt{pre-o}}$ to true, then we return true for the original proof obligation P — if the Hoare triple with an overapproximate precondition holds, then the original Hoare triple also

holds.

If the SMT solver evaluates $P_{\texttt{pre-o}}$ to false, then we submit the proof obligation $(P_{\texttt{pre-u}} : \{\phi_s^{u_{d_2}}\}(e)\{\phi_d\})$ to the SMT solver. Recall that the precondition $\phi_s^{u_{d_2}}$ is the underapproximated version of ϕ_s. If the SMT solver evaluates $P_{\texttt{pre-u}}$ to false, then we return false for the original proof obligation P — if the Hoare triple with an underapproximate precondition does not hold, then the original Hoare triple also does not hold. Further, a counterexample that falsifies $P_{\texttt{pre-u}}$ would also falsify P, and is thus usable in invariant inference and correlation procedures.

Finally, if the SMT solver evaluates $P_{\texttt{pre-u}}$ to true, then we have neither proven nor disproven P. In this case, we imprecisely (but soundly) return false for the original proof obligation P (without a counterexample). Note that both approximations of P strictly fall in Type I and are discharged as discussed in section 3.13. Revisiting our examples, the proof obligation $\{\phi_{\texttt{S2:C2}}\}(\texttt{S2} \rightarrow \texttt{S5} \rightarrow \texttt{S2}, \texttt{C2} \rightarrow \texttt{C4} \rightarrow \texttt{C2})$ $\{sum_S = sum_C\}$ is provable using a depth-1 overapproximation of the precondition $\phi_{\texttt{S2:C2}}$ — the depth-1 overapproximation retains the information that the first value in lists $\texttt{l}_S$ and $\texttt{l}_C$ are equal, and that is sufficient to prove that the new values of sum_S and sum_C are also equal (given that the old values are equal, as encoded in $\phi_{\texttt{S2:C2}}$).

Similarly, the proof obligation $\{\phi_{\texttt{S2:C2}}\}(\texttt{S2} \rightarrow \texttt{S5} \rightarrow \texttt{S2} \rightarrow \texttt{S5} \rightarrow \texttt{S2}, \texttt{C2} \rightarrow \texttt{C4} \rightarrow \texttt{C2})$ $\{sum_S = sum_C\}$ evaluates to false (with a counterexample) using a depth-2 underapproximation of the precondition $\phi_{\texttt{S2:C2}}$. In the depth-2 underapproximate version, we try to prove that if the equal lists l_S and $\texttt{Clist}_m^{\texttt{lnode}}(l_C)$ have exactly two nodes[10], then the sum of the values in the two nodes of l_S is equal to the value stored in the first node in $\texttt{l}_C$. This proof obligation will return a counterexample that maps program variables to their concrete values. We show a possible counterexample to this proof obligation below.

$$\begin{array}{l|l} \texttt{sum}_S \mapsto 3 \qquad \texttt{sum}_C \mapsto 3 & \\ \texttt{l}_S \mapsto \texttt{LCons(42,LCons(43,LNil))} & m \mapsto \left(\begin{array}{l} \texttt{0x123} \mapsto_{\texttt{lnode}} (\text{.value} \mapsto 42, \text{.next} \mapsto \texttt{0x456}), \\ \texttt{0x456} \mapsto_{\texttt{lnode}} (\text{.value} \mapsto 43, \text{.next} \mapsto \texttt{0x000}), \\ () \mapsto 77 \end{array} \right) \\ \texttt{l}_C \mapsto \texttt{0x123} & \end{array}$$

[10]The underapproximation restricts both lists to have at most two nodes; the path condition for $\texttt{S2} \rightarrow \texttt{S5} \rightarrow \texttt{S2} \rightarrow \texttt{S5} \rightarrow \texttt{S2}$ additionally restricts l_S to have at least two nodes; together, this is equivalent to the list having exactly two nodes

This counterexample maps variables to values (e.g., $\texttt{sum}_C$ maps to an `i32` value `3` and $\texttt{l}_S$ maps to a `List` value `LCons(42,LCons(43,LNil))`). It also maps the C program's memory state m to an array that maps the regions starting at addresses `0x123` and `0x456` (regions of size '`sizeof lnode`') to memory objects of type `lnode` (with the `value` and `next` fields shown for each object). For all other addresses (except the ones for which an explicit mapping is available), m maps them to the default byte-value `77` (shown as `()` $\mapsto$ `77`) in this counterexample.

This counterexample satisfies the preconditions $l_S \approx_2 \texttt{Clist}_m^{\texttt{lnode}}(\texttt{l}_C)$ and $sum_S = sum_C$. Further, when the paths $(\texttt{S2} \to \texttt{S5} \to \texttt{S2} \to \texttt{S5} \to \texttt{S2}, \texttt{C2} \to \texttt{C4} \to \texttt{C2})$ are executed starting at the machine state represented by this counterexample, the resulting values of sum_S and sum_C are `3+42+43=88` and `3+42=45` respectively. Evidently, the counterexample falsifies the proof condition because these values are not equal (as required by the postcondition).

3.15 Type III: k-unrolling P contains recursive relations in the RHS

In fig. 11, consider a proof obligation generated across the product-CFG edge $(\texttt{S3} : \texttt{C5}) \to (\texttt{S3} : \texttt{C3})$ while checking if the (I4) invariant, $l_S \sim Clist_m^{lnode}(l_C)$, holds at `(S3:C3)`: $\{\phi_{\texttt{S3:C5}}\}(\texttt{S3} \to \texttt{S5} \to \texttt{S3}, \texttt{C5} \to \texttt{C3})\{l_S \sim Clist_m^{lnode}(l_C)\}$. Here, a recursive relation is present both in the precondition $\phi_{\texttt{S3:C5}}$ ((I8)) and in the postcondition ((I4)) and we are unable to remove them after k-unrolling. When lowered to first-order logic through $\texttt{WP}_{\texttt{S3}\to\texttt{S5}\to\texttt{S3},\texttt{C5}\to\texttt{C3}}$, this translates to (showing only relevant relations):

$$(\texttt{i}_S = \texttt{i}_C \wedge \texttt{p}_C = \texttt{malloc}() \wedge \texttt{l}_S \sim \texttt{Clist}_m^{\texttt{lnode}}(\texttt{l}_C)) \Rightarrow (\texttt{LCons}(\texttt{i}_S, \texttt{l}_S) \sim \texttt{Clist}_{m'}^{\texttt{lnode}}(\texttt{p}_C)) \tag{12}$$

On the `RHS` of this first-order logic formula, $\texttt{LCons(i}_S\texttt{, l}_S\texttt{)}$ is compared for equality with $\texttt{Clist}_{m'}^{\texttt{lnode}}\texttt{(p}_C\texttt{)}$; here $\texttt{p}_C$ represents the address of the newly allocated `lnode` object (through `malloc`) and m' represents the C memory state after executing the writes at lines `C5` and `C6` on the path `C5`→`C3`, i.e.,

$$m' \equiv m[\&(\texttt{p}_C \xrightarrow{m}_{\texttt{lnode}} \texttt{value}) \leftarrow \texttt{i}_C]_{\texttt{i32}}[\&(\texttt{p}_C \xrightarrow{m}_{\texttt{lnode}} \texttt{next}) \leftarrow \texttt{l}_C]_{\texttt{i32}} \tag{13}$$

Here, $m[\mathtt{a} \leftarrow \mathtt{v}]$ reprsents an array that is equal to m everywhere except at address $\mathtt{a}$ which contains the value $\mathtt{v}$. We also refer to these memory writes that distinguish m from m', the *distinguishing writes*.

3.15.1 LHS-to-RHS Substitution and RHS Decomposition

S2C utilizes the $\sim$ relationships in the `LHS` (antecedent) of "$\Rightarrow$" to rewrite eq. (12) so that the recursive `List` values in its `RHS` (consequent) are substituted with the lifted C values (lifted using the `Clist` constructor). Thus, we rewrite eq. (12) to:

$$(\mathtt{i}_S = \mathtt{i}_C \wedge \mathtt{p}_C = \mathtt{malloc}() \wedge \mathtt{l}_S \sim \mathtt{Clist}_m^{\mathtt{lnode}}(\mathtt{l}_C)) \Rightarrow (\mathtt{LCons}(\mathtt{i}_S, \mathtt{Clist}_m^{\mathtt{lnode}}(\mathtt{l}_C)) \sim \mathtt{Clist}_{m'}^{\mathtt{lnode}}(\mathtt{p}_C)) \tag{14}$$

Next, we decompose the `RHS` by decomposing the recursive relation in the `RHS` followed by RHS-breaking. This process reduces eq. (12) into the following smaller proof obligations (showing only the `RHS`, the `LHS` is the same as in eq. (12)): (1) $\neg(\mathtt{p}_C = 0)$, (2) $\neg(\mathtt{p}_C = 0) \rightarrow \mathtt{i}_S = (\mathtt{p}_C \xrightarrow{m'}_{\mathtt{lnode}} \mathtt{value})$, and (3) $\neg(\mathtt{p}_C = 0) \rightarrow \mathtt{Clist}_m^{\mathtt{lnode}}(\mathtt{l}_C) \sim \mathtt{Clist}_{m'}^{\mathtt{lnode}}(\mathtt{p}_C \xrightarrow{m'}_{\mathtt{lnode}} \mathtt{next})$. The first two proof obligations fall in *Type II* and are discharged through over- and under-approximation schemes (as discussed in section 3.14):

1. The first proof obligation with postcondition $\neg(\mathtt{p}_C = 0)$ evaluates to true because the `LHS` ensures that $\mathtt{p}_C$ is the return value of an allocation function (`malloc`) which must be non-null due to the (C `fits`) assumption.

2. The second proof obligation with postcondition $(\mathtt{i}_S = (\mathtt{p}_C \xrightarrow{m'}_{\mathtt{lnode}} \mathtt{value}))$ also evaluates to true because $\mathtt{i}_C$ is written to address $\&(\mathtt{p}_C \xrightarrow{m'}_{\mathtt{lnode}} \mathtt{value})$ in m' (eq. (13)) and the `LHS` ensures that $\mathtt{i}_S = \mathtt{i}_C$.

For ease of exposition, we simplify the postcondition of the third proof obligation from $\neg(\mathtt{p}_C = 0) \rightarrow (\mathtt{Clist}_m^{\mathtt{lnode}}(\mathtt{l}_C) \sim \mathtt{Clist}_{m'}^{\mathtt{lnode}}(\mathtt{p}_C \xrightarrow{m'}_{\mathtt{lnode}} \mathtt{next}))$ to $(\mathtt{Clist}_m^{\mathtt{lnode}}(\mathtt{l}_C) \sim \mathtt{Clist}_{m'}^{\mathtt{lnode}}(\mathtt{l}_C))$. This simplification is valid because $\mathtt{l}_C$ is written to address $\&(\mathtt{p}_C \xrightarrow{m'}_{\mathtt{lnode}} \mathtt{next})$ in m' (eq. (13)). Also, we have already shown that $\neg(\mathtt{p}_C = 0)$ holds. Thus, the third proof obligation can be rewritten as a recursive relation between two lifted expressions[11]:

$$\mathtt{Clist}_m^{\mathtt{lnode}}(\mathtt{l}_C) \sim \mathtt{Clist}_{m'}^{\mathtt{lnode}}(\mathtt{l}_C) \tag{15}$$

[11]This simplification-based rewriting is only shown for ease of exposition, and has no effect on the operation of the algorithm. Even if the proof obligation is not simplified, the unification-

Hence, we are interested in proving equality between two `List` values in C under different memory states m and m'. Next, we show how the above can be reformulated to the problem of showing equivalence between two procedures through bisimulation.

3.15.2 Conversion of recursive equality between lifted expressions to a bisimulation

Consider a program that recursively calls the unrolling procedure in eq. (9) to deconstruct $\texttt{Clist}_m^{\texttt{lnode}}(\texttt{l}_C)$. For example, $\texttt{Clist}_m^{\texttt{lnode}}(\texttt{l}_C)$ may yield a recursive call to the unrolling procedure $\texttt{Clist}_m^{\texttt{lnode}}(\texttt{l}_C \xrightarrow{m}_{\texttt{lnode}} \texttt{next})$ and so on, until the argument to the unrolling procedure becomes zero. This program essentially deconstructs $\texttt{Clist}_m^{\texttt{lnode}}(\texttt{l}_C)$ into its terminal (scalar) values and reconstructs a `List` value equal to the value represented by $\texttt{Clist}_m^{\texttt{lnode}}(\texttt{l}_C)$. We call this program a *reconstruction program* based on the unrolling procedure of $\texttt{Clist}_m^{\texttt{lnode}}(\texttt{l}_C)$.

Theorem 1. *Under the antecedent* $(\texttt{l}_S \sim \texttt{Clist}_m^{\texttt{lnode}}(\texttt{l}_C))$*:*

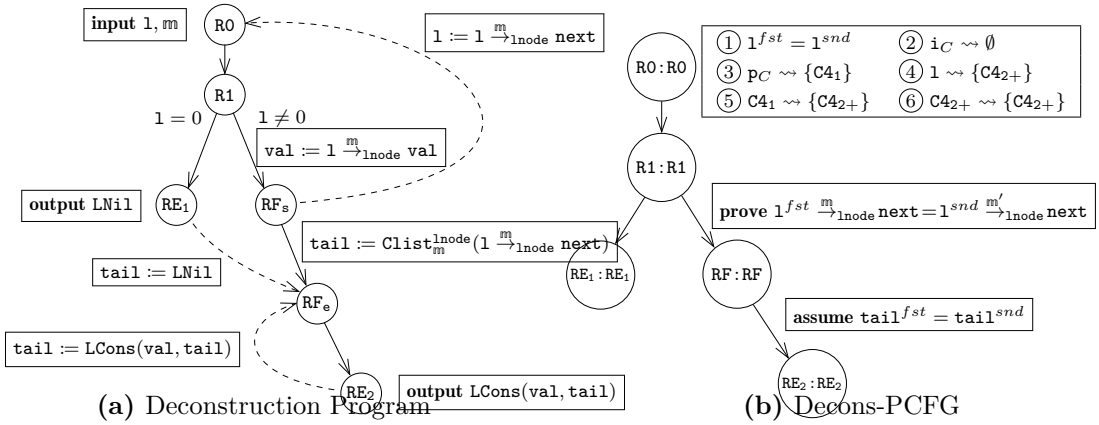
$(\texttt{Clist}_m^{\texttt{lnode}}(\texttt{l}_C) \sim \texttt{Clist}_{m'}^{\texttt{lnode}}(\texttt{l}_C))$ *holds iff a bisimulation relation exists between the reconstruction programs based on* $\texttt{Clist}_m^{\texttt{lnode}}(\texttt{l}_C)$ *and* $\texttt{Clist}_{m'}^{\texttt{lnode}}(\texttt{l}_C)$*. The bisimulation relation must ensure that the observables generated by both procedures are identical.*

Proof. The "if" case of this "iff" relation follows from noting that the observables of a reconstruction program are the generated `List` values. Thus, a successful bisimulation check ensures equal `List` values upon termination. Termination follows from the antecedent because Spec values (such as l_S) must be finite.

The "only if" case follows from the unification of the unrolling procedure (in eq. (9)) for $\texttt{Clist}_m^{\texttt{lnode}}(\texttt{l}_C)$ and $\texttt{Clist}_{m'}^{\texttt{lnode}}(\texttt{l}_C)$.

Thus, to check if $\texttt{Clist}_m^{\texttt{lnode}}(\texttt{l}_C) \sim \texttt{Clist}_{m'}^{\texttt{lnode}}(\texttt{l}_C)$, we check if a bisimulation exists between the two respective reconstruction programs (potentially under an

based proof discharge algorithm will generate proof conditions of the form $\neg(\texttt{p}_C = 0) \Rightarrow ((\texttt{p}_C \xrightarrow{m'}_{\texttt{lnode}}\texttt{next})=\texttt{l}_C)$ which will be successfully discharged by the SMT solver.



(a) Deconstruction Program **(b)** Decons-PCFG

Figure 14: The deconstruction program for $\texttt{Clist}_m^{\texttt{lnode}}(\texttt{l}_C)$ and decons-PCFG between deconstruction programs of $\texttt{Clist}_m^{\texttt{lnode}}(\texttt{l}_C)$ and $\texttt{Clist}_{m'}^{\texttt{lnode}}(\texttt{l}_C)$. In fig. 14a, `D0` represents the unrolling procedure entry node, and the square boxes show the transfer functions of the unrolling procedure (eq. (9)). The dashed edges represent a recursive function call. In fig. 14b, the square box to the right of node `D0:D0` contains the inferred invariants for this decons-PCFG.

antecedent). Theorem 1 generalizes to equality of arbitrary lifted expressions constructed from potentially different C values and memory states.

3.15.3 Checking bisimulation between reconstruction programs

To check bisimulation, we attempt to show that both reconstructions proceed in lockstep, and the invariants at each step of this lockstep execution ensure equal observables. We use a product-CFG to encode this lockstep execution — to distinguish this product-CFG from the top-level product-CFG that relates S and C, we call this product-CFG that relates two reconstruction programs, a *reconstruction product-CFG* or *recons-PCFG* for short.

The reconstruction program and the recons-PCFG for our `Clist` example are shown in fig. 14. To check bisimulation between the programs that deconstruct $\texttt{Clist}_m^{\texttt{lnode}}(\texttt{l}_C)$ and $\texttt{Clist}_{m'}^{\texttt{lnode}}(\texttt{l}_C)$, the recons-PCFG correlates one unrolling of the first program with one unrolling of the second program. An unrolling of each reconstruction program is based on the unrolling procedure in eq. (9). Thus, the PC-transition correlations of both programs are trivially obtained by unifying the unrolling procedure with itself. A node is created in the recons-PCFG that encodes the correlation of the entries of the unrolling procedure in both programs,

we call this node the *recursive-node* in the recons-PCFG, e.g., the recursive node in fig. 14b is `R0:R0`. A recursive call becomes a back-edge in the recons-PCFG that terminates at the recursive-node. A candidate invariant at the recursive-node is obtained by equating the pair of corresponding $\texttt{l}_C$ variables across the first and second programs, i.e., $\texttt{l}_C^{fst} = \texttt{l}_C^{snd}$. At the start of both reconstruction programs, $\texttt{l}_C^{fst} = \texttt{l}_C^{snd} = \texttt{l}_C^{start}$ — the same $\texttt{l}_C^{start}$ is passed to both reconstruction programs, only the memory states m and m' are different. The bisimulation check thus involves checking that if the invariant $\texttt{l}_C^{fst} = \texttt{l}_C^{snd}$ holds at the recursive-node, then during one iteration of the unrolling procedure in both programs:

1. The if condition $(\texttt{l}_C^{fst} = 0)$ in the first program is equal to the corresponding if condition $(\texttt{l}_C^{snd} = 0)$ in the second program.

2. If the if condition evaluates to false in both programs, then the observable values (that are used in the construction of the list) are equal, i.e., $((\texttt{l}_C^{fst} \neq 0) \wedge (\texttt{l}_C^{snd} \neq 0)) \Rightarrow (\texttt{l}_C^{fst} \xrightarrow{m}_{\texttt{lnode}} \texttt{val} = \texttt{l}_C^{snd} \xrightarrow{m'}_{\texttt{lnode}} \texttt{val})$.

3. If the if condition evaluates to false in both programs, then the invariant holds at the beginning of the unrolling procedure invoked through the recursive call. This involves checking equality of the arguments to the recursive call, i.e., $((\texttt{l}_C^{fst} \neq 0) \wedge (\texttt{l}_C^{snd} \neq 0)) \Rightarrow \texttt{l}_C^{fst} \xrightarrow{m}_{\texttt{lnode}} \texttt{next} = \texttt{l}_C^{snd} \xrightarrow{m'}_{\texttt{lnode}} \texttt{next}$.

The first check succeeds due to the invariant $\texttt{l}_C^{fst} = \texttt{l}_C^{snd}$. For the second and third checks, we additionally need to reason that the memory objects $\texttt{l}_C^{fst} \xrightarrow{m}_{\texttt{lnode}} \texttt{val}$ and $\texttt{l}_C^{fst} \xrightarrow{m}_{\texttt{lnode}} \texttt{next}$ cannot alias with the writes (in m' in eq. (13)) to the newly allocated objects $\texttt{p}_C \xrightarrow{m}_{\texttt{lnode}} \texttt{val}$ and $\texttt{p}_C \xrightarrow{m}_{\texttt{lnode}} \texttt{next}$. This aliasing information is captured using a points-to analysis, described next in section 3.15.4.

Notice that a bisimulation check between the reconstruction programs is significantly easier than the top-level bisimulation check between Spec and C programs: here, the correlation of PC transitions is trivially identified by unifying the unrolling procedure with itself, and the candidate invariants are obtained by equating each corresponding pair of variables across the two programs.

3.15.4 Points-to Analysis

To reason about aliasing (as required during the bisimulation check in section 3.15.3), we conservatively compute the *may-point-to* information for each program value using Andersen's algorithm [10]. The range of this computed may-point-to function are *sets of region labels*, where each region label identifies a set of memory objects. The sets of memory objects identified by two distinct region labels are necessarily disjoint. We write $p \rightsquigarrow \{R_1, R_2\}$ to represent the condition that value p may point to an object belonging to one of the region labels R_1 or R_2 (but may not point to any object outside of R_1 and R_2).

We populate the set of all region labels using the *allocation sites* of the program, i.e., PCs where a call to `malloc` exists, e.g., `C4` in fig. 2b is an allocation site. For each allocation site A, we create two region labels: (1) the first region label, called A_1, identifies the set of memory objects that were allocated by the most recent execution of A. (2) The second region label, called A_{2+}, identifies the set of memory objects that were allocated by older (not the most recent) executions of A.

For example, at the start of PC `C7` in fig. 2b, $\mathtt{i}_C \rightsquigarrow \emptyset$, $\mathtt{n}_C \rightsquigarrow \{\mathtt{C4}_1\}$, and $\mathtt{l}_C \rightsquigarrow \{\mathtt{C4}_{2+}\}$. Because the may-point-to analysis determines the sets of objects pointed-to by $\mathtt{n}_C$ and $\mathtt{l}_C$ to be disjoint ($\{\mathtt{C4}_1\}$ vs. $\{\mathtt{C4}_{2+}\}$), any memory accesses through $\mathtt{n}_C$ and $\mathtt{l}_C$ cannot alias at `C7` (for an access offset that is within the bounds of the allocation size '`sizeof lnode`').

The may-point-to information is computed not just for scalar program values ($\mathtt{n}_C$, $\mathtt{l}_C$, ...) but also for each region label. For region labels $A1_{r1}$, $A2_{r2}$, $A3_{r3}$: $A1_{r1} \rightsquigarrow \{A2_{r2}, A3_{r3}\}$ represents the condition that the values (pointers) stored in objects identified by $A1_{r1}$ may point to an object identified by either $A2_{r2}$ or $A3_{r3}$ (but not to any object outside $A2_{r2}$ and $A3_{r3}$). In fig. 2b, at PC `C7`, we get $\mathtt{C4}_1 \rightsquigarrow \{\mathtt{C4}_{2+}\}$ and $\mathtt{C4}_{2+} \rightsquigarrow \{\mathtt{C4}_{2+}\}$. The condition $\mathtt{C4}_1 \rightsquigarrow \{\mathtt{C4}_{2+}\}$ holds because the `next` pointer of the object pointed-to by $\mathtt{n}_C$ (which is a $\mathtt{C4}_1$ object) may point to a $\mathtt{C4}_{2+}$ object (e.g., object pointed-to by $\mathtt{l}_C$). Similarly, $\mathtt{C4}_{2+} \rightsquigarrow \{\mathtt{C4}_{2+}\}$ says that a pointer within a $\mathtt{C4}_{2+}$ object may point to a $\mathtt{C4}_{2+}$ object (but not to a $\mathtt{C4}_1$ object).

3.15.5 Transferring points-to information to the recons-PCFG

Recall that in section 3.15.2, we reduce a validity check of the condition $\texttt{Clist}_m^{\texttt{lnode}}(\texttt{l}_C) \sim \texttt{Clist}_{m'}^{\texttt{lnode}}(\texttt{l}_C)$ to a bisimulation check. Also, recall that we discharge the bisimulation check through the construction of a recons-PCFG that compares the unrolling procedure with itself (executing on memory states m and m'). During this bisimulation check, we need to prove that for each execution of the unrolling procedure, $\texttt{l}_C \xrightarrow{m}_{\texttt{lnode}} \{\texttt{val}, \texttt{next}\}$ and $\texttt{l}_C \xrightarrow{m'}_{\texttt{lnode}} \{\texttt{val}, \texttt{next}\}$[12] are equal. To successfully discharge these proof obligations, it suffices to show $\texttt{l}_C$ cannot alias with the memory writes that distinguish m from m'.

Our points-to analysis on the C program determines that at PC `C5` (the start of the product-CFG edge `(S3:C5)`→`(S3:C3)` across which the proof condition is being evaluated), the pointer to the *head* of the list, i.e., $\texttt{l}_C^{start}$ points to $\texttt{C4}_{2+}$. It also determines that the distinguishing writes modify memory regions belonging to $\texttt{C4}_1$. Further, we get $\texttt{C4}_{2+} \rightsquigarrow \{\texttt{C4}_{2+}\}$ at PC `C5`. However, notice that these determinations only rule out aliasing of the list-head with the distinguishing writes. We also need to confirm non-aliasing of the internal nodes of the linked list with the distinguishing writes. For this, we need to identify a points-to invariant, $\texttt{l}_C \rightsquigarrow \{\texttt{C4}_{2+}\}$, at the recursive-node of the recons-PCFG (shown in fig. 14b). To see why $\texttt{l}_C \rightsquigarrow \{\texttt{C4}_{2+}\}$ is an inductive invariant at the recursive-node:

- (Base case) The invariant holds at entry to the recons-PCFG (because it holds for $\texttt{l}_C^{start}$).
- (Induction step) If $\texttt{l}_C \rightsquigarrow \{\texttt{C4}_{2+}\}$ holds at the start of an unrolling procedure, it also holds at the start of a recursive call to the unrolling procedure. This follows from $\texttt{C4}_{2+} \rightsquigarrow \{\texttt{C4}_{2+}\}$ (points-to information at PC `C5`), which ensures that $\texttt{l}_C$`->next` may point to only $\texttt{C4}_{2+}$ objects.

To identify this points-to invariant, we run our points-to analysis (the same analysis that is run on the C program) on the reconstruction programs (fig. 14a) before comparing them for equivalence. The boundary condition for the points-to analysis at the entry node of the reconstruction program (e.g., `R0` in fig. 14) is based on the results of the points-to analysis on C at the PC where the proof obligation

[12]Here, we use the symbol $\texttt{l}_C$ to refer to equal values $\texttt{l}_C^{fst}$ and $\texttt{l}_C^{snd}$.

is being discharged (e.g., C5 in our fig. 1b). The points-to invariants at a node $(\mathtt{R}_i^{fst}, \mathtt{R}_j^{snd})$ of a recons-PCFG are derived from the results of the points-to analysis on the individual reconstruction programs at nodes $\mathtt{R}_i^{fst}$ and $\mathtt{R}_j^{snd}$ respectively.

During proof obligation discharge (e.g., during the bisimulation check on recons-PCFG), the points-to invariants are encoded as SMT constraints. This allows us to successfully complete the bisimulation proof on the recons-PCFG, and consequently successfully discharge the proof obligation $\{\phi_{\mathtt{S3:C5}}\}$
$(\mathtt{S3} \rightarrow \mathtt{S5} \rightarrow \mathtt{S3}, \mathtt{C5} \rightarrow \mathtt{C3})\{l_S \sim Clist_m^{\mathtt{lnode}}(l_C)\}$ in table 2. The points-to analysis is described more formally in section 4.5.

3.15.6 Proof discharge algorithm for Type III obligations

Before the start of an equivalence check, a points-to analysis is run on the C IR once. During the equivalence check, to discharge a Type III proof obligation $P : \mathtt{LHS} \Rightarrow \mathtt{RHS}$ (expressed in first-order logic), we first replace the recursive values of program S in the RHS with lifted C values, based on the equalities present in the LHS, to obtain P_2. This is followed by decomposition and RHS-breaking of P_2.

Upon successful decomposition, we obtain several smaller proof obligations. To prove P, we require all these smaller proof obligations to be provable. If any of these smaller proof obligations is not provable, we are unable to prove P. If we obtain a counterexample to any of these smaller proof obligations, then that counterexample also falsifies P. Let P_3 represent any such smaller proof obligation. RHS of P_3, being a decomposition clause, must relate atomic expressions on the RHS. If P_3 relates two scalar values in the RHS, then it is a Type II proof obligation and can be discharged using the algorithm in section 3.14.4.

If P_3 relates two lifted expressions in the RHS, we check if the reconstruction programs of the two lifted ADT values being compared can be proven to be bisimilar (assuming that LHS of P_3 holds at the correlated entry nodes in the recons-PCFG). To improve the bisimulation check's precision, we transfer the points-to information of the C program (at the PC where the proof obligation is being discharged) to the entry of the reconstruction programs. The same points-to analysis is ran on the reconstruction programs to populate the points-to function at all PCs.

These queries generated by a bisimulation check are discharged by a recursive call to the proof discharge procedure. The depth of these recursive calls to the

proof discharge procedure is determined by the maximum *recursion nest depth* (similar to loop nest depth) of the decomposition program.

If the bisimilarity check succeeds, the proof procedure returns true for P. If the bisimilarity check fails, we imprecisely return false for P (without a counterexample).

Finally, if P_3 neither relates two scalar values, nor relates two lifted expressions, we attempt to prove that `LHS` of P_3 imply `false`. If successfully disproven, we return false for P with the counterexamples. Otherwise, we imprecisely return false for P (without a counterexample).

Please refer to `Chapter XXX` of the thesis for a detailed discussion on the algorithms introduced in this section along with their pseudo-code.

4 Formalism

4.1 The Spec Language

We briefly discuss the properties of the Spec language in this section. Spec supports recursive algebraic data types (ADT) similar to the ones available in most functional languages. The types in Spec can be represented in *first order recursive types* with `Product` and `Sum` type constructors and `Unit, Bool, i<N>` types (i.e., nullary type constructors) as follows:

$T \rightarrow \mu\alpha \mid \texttt{Product}(T, \ldots, T) \mid \texttt{Sum}(T, \ldots, T) \mid \texttt{Unit} \mid \texttt{Bool} \mid \texttt{i}\langle\texttt{N}\rangle \mid \alpha$

For example, the `List` type can be written as $\mu\alpha.Sum(Unit, Prod(i32, \alpha))$.

The language also borrows its expression grammar heavily from functional languages. This includes the usual constructs like `let-in`, `if-then-else`, function application and the `match` statement for pattern-matching (i.e. deconstructing) sum and product values. Unlike functional languages, Spec only supports first order functions. Also, Spec does not support partial function application. Hence, we constrain our attention to C programs containing only first order functions. Spec is equipped with a special `assuming-do` construct for explicitly providing UB conditions. These assumptions become part of (S `def`) as discussed in section 2.3.

Spec also provides the typical boolean and bitvector operators for expressing computation in C succintly yet explicitly. This includes logical operators (e.g., `and`), bitvector arithmatic operators (e.g., `bvadd(+)`) and relational operators for comparing bitvectors interpreted as signed or unsigned integers (e.g., $\leq_{u,s}$).

$$
\begin{array}{rcl}
\langle \text{expr} \rangle & \rightarrow & \texttt{if}\ \langle \text{expr} \rangle\ \texttt{then}\ \langle \text{expr} \rangle\ \texttt{else}\ \langle \text{expr} \rangle \\
& | & \texttt{let}\ \langle \text{id} \rangle\ \texttt{=}\ \langle \text{expr} \rangle\ \texttt{in}\ \langle \text{expr} \rangle \\
& | & \texttt{match}\ \langle \text{expr} \rangle\ \texttt{with}\ \langle \text{match-clause-list} \rangle \\
& | & \texttt{assuming}\ \langle \text{expr} \rangle\ \texttt{do}\ \langle \text{expr} \rangle \\
& | & \langle \text{id} \rangle\ \texttt{(}\ \langle \text{expr-list} \rangle\ \texttt{)} \\
& | & \langle \text{data-cons} \rangle\ \texttt{(}\ \langle \text{expr-list} \rangle\ \texttt{)} \\
& | & \langle \text{expr} \rangle\ \texttt{is}\ \langle \text{data-cons} \rangle \\
& | & \langle \text{expr} \rangle\ \langle \text{scalar-op} \rangle\ \langle \text{expr} \rangle \\
& | & \langle \text{literal}_{\text{Unit}} \rangle \mid \langle \text{literal}_{\text{Bool}} \rangle \mid \langle \text{literal}_{\text{i<N>}} \rangle \\
\\
\langle \text{match-clause-list} \rangle & \rightarrow & \langle \text{match-clause} \rangle^* \\
\langle \text{match-clause} \rangle & \rightarrow & \mid \langle \text{data-cons} \rangle\ \texttt{(}\ \langle \text{id-list} \rangle\ \texttt{)} \Rightarrow \langle \text{expr} \rangle \\
\langle \text{expr-list} \rangle & \rightarrow & \epsilon \mid \langle \text{expr} \rangle\ \texttt{,}\ \langle \text{expr-list} \rangle \\
\langle \text{id-list} \rangle & \rightarrow & \epsilon \mid \langle \text{id} \rangle\ \texttt{,}\ \langle \text{id-list} \rangle \\
\\
\langle \text{literal}_{\text{Unit}} \rangle & \rightarrow & \texttt{()} \\
\langle \text{literal}_{\text{Bool}} \rangle & \rightarrow & \texttt{false} \mid \texttt{true} \\
\langle \text{literal}_{\text{i<N>}} \rangle & \rightarrow & [\texttt{0}\ldots\texttt{2}^{\text{N}}\texttt{-1}]
\end{array}
$$

Figure 15: Simplified expression grammar of Spec language

4.2 Counterexample-guided Best-First Search Algorithm for a Product-CFG

S2C constructs a product-CFG incrementally to search for an observably-equivalent bisimulation relation between the individual CFGs of a Spec program S and a C program C. Multiple candidate product-CFGs are partially constructed during this search; the search completes when one of these candidates yields an equivalence proof.

Anchor nodes in the CFG of the C program are identified to ensure that every cycle in the CFG contains at least one anchor node. Also, for every procedure call in the CFG, anchor nodes are created just before and just after the callsite, e.g., in fig. 10b, `C4` and `C5` are anchor nodes around the call to `malloc()`. Our algorithm ensures that for each anchor node in C, we identify a correlated node in S — if

a product-CFG π contains a product-CFG node (n_S, n_C), then π correlates node n_C in C with node n_S in S. The first partially-constructed product-CFG contains a single entry node that encodes the correlation of the entry nodes (`S0:C0`) of the two input CFGs.

At each step of the incremental construction algorithm, a node (n_S, n_C) is chosen in a product-CFG π and a path ρ_C in C's CFG starting at n_C (and ending at an anchor node in C) is selected. Then, the potential correlations ρ_C with paths in S's CFG are enumerated. For example, in fig. 11, at product-CFG node `(S3:C3)`, we first select the C path `C3`→`C4`, and its potential correlation possibilities with paths ϵ, `S3` → `S5`, `S3` → `S5` → `S3`, `S3` → `S5` → `S3` → `S5`, ... in S are enumerated (up to an unroll factor μ).

For each enumerated correlation possibility (ρ_S, ρ_C), a separate product-CFG π' is created (by cloning π) and a new product-CFG edge $e = (\rho_S, \rho_C)$ is added to π'. The head of the product-CFG edge e is the (potentially newly added) product-CFG node representing the correlation of the end-points of paths ρ_S and ρ_C. For example, the node `(S3:C4)` is added to the product-CFG if it correlates paths ϵ and `C3` → `C4` starting at `(S3:C3)`. For each node s in a product-CFG π, we maintain a small number of concrete machine state pairs (of S and C) at s. The concrete machine state pairs at s are obtained as counterexamples to an unsucessful proof obligation $\{\phi_s\}(s \to d)\{\phi_d\}$ (for some edge $s \to d$ and node d in π). Thus, by construction, these counterexamples represent concrete state pairs that may potentially occur at s during the lockstep execution encoded by π.

To evaluate the promise of a possible correlation (ρ_S, ρ_C) starting at node s in product-CFG π, we examine the execution behavior of the counterexamples at s on the product-CFG edge $e = (s \to d) = (\rho_S, \rho_C)$. If the counterexamples ensure that the machine states remain related at d, then that candidate correlation is ranked higher. This ranking criterion is based on prior work [22]. A best-first search (BFS) procedure based on this ranking criterion is used to incrementally construct a product-CFG that proves bisimulation. For each intermediate candidate product-CFG π generated during this search procedure, an automatic invariant inference procedure is used to identify invariants at all the nodes in π. The counterexamples obtained from the proof obligations created by this invariant inference procedure are added to the respective nodes in π; these counterexamples help rank future

correlations starting at those nodes.

If after invariant inference, we realize that an intermediate candidate product-CFG π_1 is not promising enough, we backtrack and choose another candidate product-CFG π_2 and explore the potential correlations that can be added to π_2. Thus, a product-CFG is constructed one edge at a time. If at any stage, the invariants inferred for a product-CFG π_i ensure equal observables, we have successfully demonstrated equivalence.

This counterexample-guided BFS procedure is similar to the one described in prior work on the Counter algorithm [22]. Our primary contribution is a proof discharge algorithm for proof obligations containing recursive relations (sections 3.9, 3.10 and 3.13 to 3.15). These proof obligations may be generated either at the intermediate (search) or the final (check) phases of the BFS procedure.

Table 3: Dataflow formulation for the Invariant Inference Algorithm.

Domain	$\left\{\begin{array}{c}\phi_n \text{ is a conjunction of predicates drawn from} \\ \text{grammar in 16b, } \Gamma_n \text{ is a set of counterexamples}\end{array}\right\}$
Direction	Forward
Transfer function across edge $e = (s \to d)$	$(\phi_d, \Gamma_d) = f_e(\phi_s, \Gamma_s)$ (fig. 16a)
Meet operator $\otimes$ $(\phi_n, \Gamma_n) \leftarrow (\phi_n^1, \Gamma_n^1) \otimes (\phi_n^2, \Gamma_n^2)$	$\Gamma_n \leftarrow \Gamma_n^1 \cup \Gamma_n^2$, $\quad \phi_n \leftarrow \mathit{StrongestInvCover}(\Gamma_n)$
Boundary condition	$\texttt{out}[n^{start}] = (Pre, \Gamma_{n^{start}})$
Initialization to $\top$	$\texttt{in}[n] = (\texttt{False}, \{\})$ for all non-start nodes

Function $f_e(\phi_s, \Gamma_s)$
  $\Gamma_d^{can} \coloneqq \Gamma_d \cup \texttt{exec}_e(\Gamma_s)$;
  $\phi_d^{can} \coloneqq \mathit{StrongestInvCover}(\Gamma_d^{can})$;
  while $\texttt{SAT}(\neg(\{\phi_s\}(e)\{\phi_d^{can}\}), \gamma_s)$ **do**
    $\gamma_d \coloneqq \texttt{exec}_e(\gamma_s)$;
    $\Gamma_d^{can} \coloneqq \Gamma_d^{can} \cup \gamma_d$;
    $\phi_d^{can} \coloneqq \mathit{StrongestInvCover}(\Gamma_d^{can})$;
  end
  return $(\phi_d^{can}, \Gamma_d^{can})$;
end

(a) Transfer function f_e across edge $e = (s \to d)$.

$$Inv \;\to\; \textstyle\sum_i c_i v_i = c \mid v_1 \odot v_2 \mid \alpha_S = \texttt{liftC}_m(v^C, \ldots)$$

(b) Predicate grammar for constructing invariants. v represents a bitvector variable in either S or C. c represents a bitvector constant. $\odot \in \{<, \leq\}$. α_S represents an ADT variable in Spec. v^C represents a bitvector variable in C. m represents the current C memory state.

Figure 16: Transfer function f_e and Predicate grammar Inv for invariant inference dataflow analysis in table 3. Given invariants (ϕ_s) and counterexamples (Γ_s) at node s, f_e returns the updated invariants (ϕ_d) and counterexamples (Γ_d) at node d. *StrongestInvCover*(Γ) computes the strongest invariant cover for counterexamples Γ. $\texttt{exec}_e(\Gamma)$ (concretely) executes counterexamples Γ over edge e. $\texttt{SAT}(\phi, \gamma)$ determines the satisfiability of ϕ; if satisfiable, the models (counterexamples) are returned in output parameter γ.

4.3 Invariant Inference and Counterexample Generation

Table 3 presents our dataflow analysis for inferring invariants ϕ_n at each node n of a product-CFG, while also generating a set of counterexamples Γ_n at node n that represents the potential concrete machine states at n.

Given the invariants and counterexamples at node s (ϕ_s, Γ_s), the transfer function initializes the new candidate set of counterexamples at d (Γ_d^{can}) to the current set of counterexamples at d (Γ_d) union-ed with the counterexamples obtained by executing Γ_s on edge e $(\texttt{exec}_e)$. The candidate invariant at d (ϕ_d^{can}) is computed as the strongest cover of Γ_d^{can} (*StrongestInvCover()*). At each step, the transfer function attempts to prove $\{\phi_s\}(e)\{\phi_d^{can}\}$ (by checking SATisfiability of its negation). If the proof succeeds, the candidate invariant ϕ_d^{can} is returned alongwith the counterexamples Γ_d^{can} learned so far. Else the candidate invariant ϕ_d^{can} is weakened using the counterexamples obtained from the SAT query (γ) and the proof attempt is repeated.

The predicate grammar allows the automatic inference of affine and inequality relations between bitvector values of both programs, as well as, recursive relations between an ADT value in Spec (α_S) and a *lifted* ADT value from C $(\texttt{liftC}_m(p_C))$. We enumerate these recursive relation guesses for all bitvector variables v^C in C and candidate liftC lifting constructor. In our implementation, the candidate liftC constructors are derived from the constructors present in the precondition *Pre* and the postcondition *Post*. More sophisticated strategies for automatic guessing of these lifting constructors are possible.

StrongestInvCover() for affine relations involves identifying the basis vectors of the kernel of the matrix formed by the counterexamples in the bitvector domain [30, 15]. For inequality relations, *StrongestInvCover(*Γ*)* returns false iff any counterexample in Γ evaluates the relation to false — this effectively simulates the Houdini approach [21]. In case of recursive relations, *StrongestInvCover(*Γ*)* attempts to disprove the recursive relation $l_1 \sim l_2$ by evaluating its depth-η underapproximation $l_1 \sim_\eta l_2$ for each counterexample in Γ and returns false if any one of them successfully evaluates to false. η is a constant parameter of the algorithm.

4.4 Modeling Procedure Calls

A top-level procedure δ in S or C may make non-tail recursive calls, e.g., for traversing a tree data structure. Our correlation algorithm (section 4.2) ensures that the anchor nodes around such a callsite are correlated one-to-one across both programs. For example, let there be a recursive call in S at PC A_S, i.e., A_S is the callsite. Then we denote the program points just before and just after this callsite as $A_S^{\mathtt{b}}$ and $A_S^{\mathtt{a}}$ respectively. Let $\mathtt{args}_{A_S}$ represent the values of the actual arguments of this procedure call. Let $\mathtt{ret}_{A_S}$ represent the values returned by this procedure call. Similarly, for a procedure call at PC A_C in C, let $A_C^{\mathtt{b}}$, $A_C^{\mathtt{a}}$, $\mathtt{args}_{A_C}$ and $\mathtt{ret}_{A_C}$ represent the before-callsite program point, after-callsite program point, arguments and return values respectively. Our algorithm ensures that the only correlations possible in a product-CFG π for these S and C program points are $A_\pi^{\mathtt{b}}\mathtt{=(}A_S^{\mathtt{b}}\mathtt{,}A_C^{\mathtt{b}}\mathtt{)}$ and $A_\pi^{\mathtt{a}}\mathtt{=(}A_S^{\mathtt{a}}\mathtt{,}A_C^{\mathtt{a}}\mathtt{)}$.

Recall that the recursive call at A_S (or A_C) must be a call to the top-level procedure δ. We utilize the user-supplied *Pre* and *Post* conditions for δ to obtain the desired invariants at nodes $A_\pi^{\mathtt{b}}$ and $A_\pi^{\mathtt{a}}$ in the product-CFG. We require a successful proof to *ensure* that $Pre(A_S^{\mathtt{args_S}}, A_C^{\mathtt{args_C}}, m_{\mathtt{b}})$ holds at $A_\pi^{\mathtt{b}}$. Further, the proof can *assume* that $Post(A_S^{\mathtt{ret_S}}, A_C^{\mathtt{ret_C}}, m_{\mathtt{a}})$ holds at $A_\pi^{\mathtt{a}}$. Here, $m_{\mathtt{b}}$ and $m_{\mathtt{a}}$ represent the memory states in C at $A_C^{\mathtt{b}}$ and $A_C^{\mathtt{a}}$ respectively. Thus, for such recursive calls to the top-level function, we inductively prove the precondition (on the arguments of the procedure call) at $A_\pi^{\mathtt{b}}$ and assume the postcondition (on the return values of the procedure call) at $A_\pi^{\mathtt{a}}$.

4.5 Points-to Analysis

We formulate our points-to analysis as a dataflow analysis as discussed below. We first identify the set R_C of all region labels representing mutually non-overlapping regions of the C memory state m. For each call to `malloc()` at PC A, we add A_1 and A_{2+} to R_C. $R_C = \bigcup_A \{A_1, A_{2+}\} \cup \{\mathtt{heap}\}$, where `heap` represents all *other* memory regions that are not captured by the region labels associated with allocation sites.

Let S_C be the set of all scalar pseudo-registers in C's IR. We use a forward dataflow analysis to identify a may-point-to function $\Delta : (S_C \cup R_C) \mapsto 2^{R_C}$ at

each program point. For an IR instruction `x := c`, for constant c, the transfer function updates $\Delta(\texttt{x}) := \emptyset$. For instruction `x := y op z` (for some arithmetic or logical operand `op`), we update $\Delta(\texttt{x}) := \Delta(\texttt{y}) \cup \Delta(\texttt{z})$. For a load instruction `x := *y`, we update $\Delta(\texttt{x})$ to $\bigcup_{R_C \in \Delta(y)} \Delta(R_C)$. For a store instruction `*x := y`, for all $R_C \in \Delta(\texttt{x})$, we update $\Delta(R_C) := \Delta(R_C) \cup \Delta(y)$. For recursive procedure calls, a *supergraph* is created by adding control flow edges from the call-site to the procedure head (copying actual arguments to the formal arguments) and from the procedure return to the returning point of the call-site (copying returned value to the variable assigned at the callsite), e.g., in fig. 14, the dashed edges represent supergraph edges. For a malloc instruction `x := malloc`$_A$`()` (where A represents the allocation site), we perform the following steps (in order):

1. Convert all existing occurrences of A_1 to A_{2+}, i.e., for all $r \in S_C \cup R_C$, if $A_1 \in \Delta(r)$, then update $\Delta(r) := (\Delta(r) \setminus \{A_1\}) \cup \{A_{2+}\}$.
2. Update $\Delta(\texttt{x}) := \{A_1\}$
3. Update $\Delta(A_{2+}) := \Delta(A_{2+}) \cup \Delta(A_1)$.
4. Update $\Delta(A_1) := \emptyset$ (empty set).

The meet operator is set-union. For a C program C, the boundary condition at entry is given by $\Delta_C^{entry}(r) = R_C$ for all $r \in S_C \cup R_C$, where Δ_P^{pc} represents the may-point-to function for program P at PC pc.

In case of a reconstruction program R, the domain of Δ contains the pseudo-registers in C's IR (S_C) as well as any region labels (R_C). In addition to these, the domain also contains the pseudo-registers of the reconstruction program itself, say R_R. For a reconstruction program R originating from a proof obligation at a product program PC (n_S, n_C), the boundary condition is given by:

$$\Delta_R^{entry}(r) = \begin{array}{ll} \Delta_C^{n_C}(r) & \text{for all } r \in S_C \cup R_C \\ \emptyset & \text{for all } r \in R_R \end{array}$$

Hence, for a reconstruction program, we use the results of the points-to analysis on C at the PC where the proof obligation is being discharged. This is a crucial step for proving equality of C values under different C memory state as seen in section 3.15.5.

5 Evaluation

We have implemented S2C on top of the Counter tool [22]. We use *four* SMT solvers running in parallel for solving SMT proof obligations discharged by our proof discharge algorithm: `z3-4.8.7`, `z3-4.8.14` [18], `Yices2-45e38fc` [19], and `cvc4-1.7` [1]. An unroll factor of *four* is used to handle loop unrolling in the C implementation. We use a default value of *eight* for over- and under-approximation depths (d_o and d_u). The default value of our unrolling parameter k (used for categorization of proof obligations) is *five*.

S2C requires the user to provide a Spec program S (specification), a C implementation C, and a file that contains the precondition Pre and postcondition $Post$. An equivalence check requires the identification of lifting constructors to relate C values to the ADT values in Spec through recursive relations. Such relations may be required at the entry of both programs (i.e. in the precondition Pre), in the middle of both programs (i.e., in the invariants at intermediate product-CFG nodes), and at the exit of both programs (i.e., in the postcondition $Post$). Pre and $Post$ are user-specified, whereas the inductive invariants are inferred automatically by our algorithm. During invariant inference, S2C derives the candidate lifting constructors from the user-specified Pre and $Post$. More sophisticated approaches to finding lifting constructors are left as future work.

5.1 Experiments

We consider programs involving four distinct ADTs, namely, (T1) `String`, (T2) `List`, (T3) `Tree` and (T4) `Matrix`. For each Spec program specification, we consider multiple C implementations that differ in their (a) layout and representation of ADTs, and (b) algorithmic strategies. For example, a `Matrix`, in C, may be laid out in a two-dimensional array, a one-dimensional array using row or column major layouts etc. On the other hand, an optimized implementation may choose manual vectorization of an inner-most loop. Next, we consider each ADT in more detail. For each, we discuss (a) its corresponding programs, (b) C memory layouts and their lifting constructors, and (c) varying algorithmic strategies.

Table 4: String lifting constructors and their definitions.

Lifting Constructor	Definition
(T1) `Str = SInvalid | SNil | SCons(i8, Str)`	
$\texttt{Cstr}_{\mathbb{m}}^{\texttt{u8[]}}(p{:}\texttt{i32})$	$\underline{\texttt{if}}\ p = 0_{\texttt{i32}}\ \ \underline{\texttt{then}}\ \texttt{SInvalid}$ $\underline{\texttt{elif}}\ p[0_{\texttt{i32}}]_{\mathbb{m}}^{\texttt{i8}} = 0_{\texttt{i8}}\ \ \underline{\texttt{then}}\ \texttt{SNil}$ $\underline{\texttt{else}}\ \texttt{SCons}(p[0_{\texttt{i32}}]_{\mathbb{m}}^{\texttt{i8}}, \texttt{Cstr}_{\mathbb{m}}^{\texttt{u8[]}}(p + 1_{\texttt{i32}}))$
$\texttt{Cstr}_{\mathbb{m}}^{\texttt{lnode(u8)}}(p{:}\texttt{i32})$	$\underline{\texttt{if}}\ p = 0_{\texttt{i32}}\ \ \underline{\texttt{then}}\ \texttt{SInvalid}$ $\underline{\texttt{elif}}\ p \xrightarrow{\mathbb{m}}_{\texttt{lnode}} \texttt{val} = 0_{\texttt{i8}}\ \ \underline{\texttt{then}}\ \texttt{SNil}$ $\underline{\texttt{else}}\ \texttt{SCons}(p \xrightarrow{\mathbb{m}}_{\texttt{lnode}} \texttt{val}, \texttt{Cstr}_{\mathbb{m}}^{\texttt{lnode(u8)}}(p \xrightarrow{\mathbb{m}}_{\texttt{lnode}} \texttt{next}))$
$\texttt{Cstr}_{\mathbb{m}}^{\texttt{clnode(u8)}}(p{:}\texttt{i32}, i{:}\texttt{i2})$	$\underline{\texttt{if}}\ p = 0_{\texttt{i32}}\ \ \underline{\texttt{then}}\ \texttt{SInvalid}$ $\underline{\texttt{elif}}\ p \xrightarrow{\mathbb{m}}_{\texttt{lnode}} \texttt{chunk}[i]_{\mathbb{m}}^{\texttt{i8}} = 0_{\texttt{i8}}\ \ \underline{\texttt{then}}\ \texttt{SNil}$ $\underline{\texttt{else}}\ \texttt{SCons}(p \xrightarrow{\mathbb{m}}_{\texttt{lnode}} \texttt{chunk}[i]_{\mathbb{m}}^{\texttt{i8}}, \texttt{Cstr}_{\mathbb{m}}^{\texttt{clnode(u8)}}(i = 3_{\texttt{i2}} ? p \xrightarrow{\mathbb{m}}_{\texttt{clnode}} \texttt{next} : p, i + 1_{\texttt{i2}}))$

5.1.1 String

We wrote a single specification in Spec for each of the following common string library functions: `strlen`, `strchr`, `strcmp`, `strspn`, `strcspn`, and `strpbrk`. For each specification program, we took multiple C implementations of that program, drawn from popular libraries like `glibc` [3], `klibc` [4], `newlib` [7], `openbsd` [8], `uClibc` [9], `dietlibc` [2], `musl` [5], and `netbsd` [6]. Some of these libraries implement the same function in two ways: one that is optimized for code size and another that is optimized for runtime. All these library implementations use a *null character* terminated array to represent a string, and the corresponding lifting constructor is $\texttt{Cstr}_{\mathbb{m}}^{\texttt{u8[]}}$. `u<N>` represents the N-bit unsigned integer type in C. For example, `u8` represents `unsigned char` type.

Further, we implemented custom C programs for all of these functions that used linked list and *chunked linked list* data structures to represent a string. In a chunked linked list, a single list node (linked through a `next` pointer) contains a small array (chunk) of values. We use a default chunk size of four for our benchmarks. The corresponding lifting constructors are $\texttt{Cstr}_{\mathbb{m}}^{\texttt{lnode(u8)}}$ and $\texttt{Cstr}_{\mathbb{m}}^{\texttt{clnode(u8)}}$ respectively. These lifting constructors are defined in table 4. $\texttt{Cstr}_{\mathbb{m}}^{\texttt{lnode(u8)}}$ requires a single argument p representing the pointer to the list node. On the other hand, $\texttt{Cstr}_{\mathbb{m}}^{\texttt{clnode(u8)}}$ requires two arguments p and i, where p represents the pointer to the chunked linked list node and i represents the position of the initial character in the chunk.

Figure 17 shows the `strlen` specification and two vastly different *C* implementations. Figure 17b is a generic implementation using a null character terminated array to represent a string similar to a C-style string. The second implementation in fig. 17c differs from fig. 17b in the following: (a) it uses a chunked linked list data layout for the input string and (b) it uses specialized bit manipulations to identify a null character in a chunk at a time. S2C is able to automatically find a bisimulation relation for both implementations against the unaltered specification. Figure 18 shows the product-CFG and invariants for each implementation.

Lifting constructors are named based on the C data layout being lifted and the Spec ADT type of the lifted value. For example, $\texttt{Cstr}^{\texttt{u8[]}}$ represents a `String` lifting constructor for an array layout. In general, we use the following naming convention for different C data layouts: `T[]` represents an array of type `T` (e.g., `u8[]`). `lnode(T)` represents a linked list node type containing a value of type `T`. Similarly, `clnode(T)` and `tnode(T)` represent a chunked linked list and a tree node with values of type `T` respectively.

Table 5: List lifting constructors and their definitions.

Lifting Constructor	Definition
(T2) `List = LNil | LCons(i32, List)`	
$\texttt{Clist}_{\mathtt{m}}^{\texttt{u32[]}}(p\ i\ n{:}\texttt{i32})$	$\underline{\texttt{if}}\ i \geq_u n\ \ \underline{\texttt{then}}\ \texttt{LNil}$ $\underline{\texttt{else}}\ \texttt{LCons}(p[i]_{\mathtt{m}}^{\texttt{i32}}, \texttt{Clist}_{\mathtt{m}}^{\texttt{u32[]}}(p, i+1_{\texttt{i32}}, n))$
$\texttt{Clist}_{\mathtt{m}}^{\texttt{lnode(u32)}}(p{:}\texttt{i32})$	$\underline{\texttt{if}}\ p = 0_{\texttt{i32}}\ \ \underline{\texttt{then}}\ \texttt{LNil}$ $\underline{\texttt{else}}\ \texttt{LCons}(p \overset{\mathtt{m}}{\rightarrow}_{\texttt{lnode}} \texttt{val}, \texttt{Clist}_{\mathtt{m}}^{\texttt{lnode}}(p \overset{\mathtt{m}}{\rightarrow}_{\texttt{lnode}} \texttt{next}))$
$\texttt{Clist}_{\mathtt{m}}^{\texttt{clnode(u32)}}(p{:}\texttt{i32}, i{:}\texttt{i2})$	$\underline{\texttt{if}}\ p = 0_{\texttt{i32}}\ \ \underline{\texttt{then}}\ \texttt{LNil}$ $\underline{\texttt{else}}\ \texttt{LCons}(p \overset{\mathtt{m}}{\rightarrow}_{\texttt{clnode}} \texttt{chunk}[i]_{\mathtt{m}}^{\texttt{i32}}, \texttt{Clist}_{\mathtt{m}}^{\texttt{clnode}}(i = 3_{\texttt{i2}} ? p \overset{\mathtt{m}}{\rightarrow}_{\texttt{clnode}} \texttt{next} : p, i+1_{\texttt{i2}}))$

5.1.2 List

We wrote a Spec program specification that creates a list, a program that traverses a list to compute the sum of its elements and a program that computes the dot product of two lists. We use three different data layouts for a list in C: array ($\texttt{Clist}_{\mathtt{m}}^{\texttt{u32[]}}$), linked list ($\texttt{Clist}_{\mathtt{m}}^{\texttt{lnode(u32)}}$), and a chunked linked list ($\texttt{Clist}_{\mathtt{m}}^{\texttt{clnode(u32)}}$). The lifting constructors are shown in table 5. Although similar to the String lifting constructors, these lifting constructors differ widely in their data encoding. For example, $\texttt{Clist}_{\mathtt{m}}^{\texttt{u32[]}}(p, i, n)$ represents a `List` value constructed from a C array p of size n starting at the i^{th} index. The list becomes empty when we are at the end

```
S0: i32 strlen (Str s) {
S1:   i32 len := 0_i32;
S2:   while ¬(s is SNil):
S3:     assume ¬(s is SInvalid);
S4:     // (s is SCons)
S5:     s   := s.tail;
S6:     len := len + 1_i32;
S7:   return len;
SE: }
```

(a) Strlen Specification

```
   size_t strlen(char* s);
C0: i32 strlen (i32 s) {
C1:   i32 i := 0_i32;
C2:   while s[0_i32]^i8_m ≠ 0_i8:
C3:     s := s + 1_i32;
C4:     i := i + 1_i32;
C5:   return i;
CE: }
```

(b) Strlen Implementation using Array

```
   typedef struct clnode {
     char chunk[4]; struct clnode* next; } clnode;
   size_t strlen(clnode* cl);
C0 : i32 strlen (i32 cl) {
C1 :   i32 hi := 0x80808080_i32; i32 lo := 0x01010101_i32;
C2 :   i32 i  := 0_i32;
C3 :   while true:
C4 :     i32 dword_ptr := &cl →^m_clnode chunk;
C5 :     i32 dword     := dword_ptr[0_i32]^i32_m;
C6 :     if ((dword − lo) & (∼ dword) & hi) ≠ 0_i32:
C7 :       if dword_ptr[0_i32]^i8_m = 0_i8: return i;
C8 :       if dword_ptr[1_i32]^i8_m = 0_i8: return i + 1_i32;
C9 :       if dword_ptr[2_i32]^i8_m = 0_i8: return i + 2_i32;
C10:       if dword_ptr[3_i32]^i8_m = 0_i8: return i + 3_i32;
C11:     cl := cl →^m_clnode next; i  := i + 4_i32;
CE : }
```

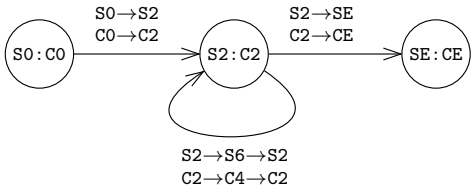
(c) Optimized Strlen Implementation using Chunked Linked List

Figure 17: Specification of Strlen along with two possible C implementations. Figure 17b is a generic implementation using a null-terminated array for `String`. Figure 17c is an optimized implementation using a chunked linked list for `String`.

of the array. $(\texttt{Clist}_{\texttt{m}}^{\texttt{lnode(u32)}})$ and $(\texttt{Clist}_{\texttt{m}}^{\texttt{clnode(u32)}})$, on the other hand, encodes empty lists (`LNil`) using *null pointers*. These layouts are in contrast to the `String` layouts, all of which uses a *null character* to indicate the empty string.

Table 6: Tree lifting constructors and their definitions.

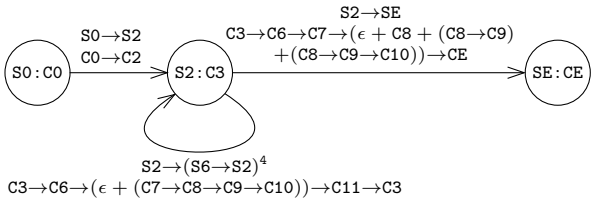
Lifting Constructor	Definition
(T3) `Tree = TNil | TCons(i32, Tree, Tree)`	
$\texttt{Ctree}_{\texttt{m}}^{\texttt{u32[]}}(p\ i\ n{:}\texttt{i32})$	$\underline{\texttt{if}}\ i \geq_u n\ \ \underline{\texttt{then}}\ \texttt{TNil}$ $\underline{\texttt{else}}\ \texttt{TCons}(p[i]^{\texttt{i32}}_{\texttt{m}}, \texttt{Ctree}_{\texttt{m}}^{\texttt{u32[]}}(p, 2_{\texttt{i32}} \times i + 1_{\texttt{i32}}, n), \texttt{Ctree}_{\texttt{m}}^{\texttt{u32[]}}(p, 2_{\texttt{i32}} \times i + 2_{\texttt{i32}}, n))$
$\texttt{Ctree}_{\texttt{m}}^{\texttt{tnode(u32)}}(p{:}\texttt{i32})$	$\underline{\texttt{if}}\ p = 0_{\texttt{i32}}\ \ \underline{\texttt{then}}\ \texttt{TNil}$ $\underline{\texttt{else}}\ \texttt{TCons}(p \overset{\texttt{m}}{\rightarrow}_{\texttt{tnode}} \texttt{val}, \texttt{Ctree}_{\texttt{m}}^{\texttt{tnode(u32)}}(p \overset{\texttt{m}}{\rightarrow}_{\texttt{tnode}} \texttt{left}), \texttt{Ctree}_{\texttt{m}}^{\texttt{tnode(u32)}}(p \overset{\texttt{m}}{\rightarrow}_{\texttt{tnode}} \texttt{right}))$



(a) Product CFG for programs figs. 17a and 17b

PC-Pair	Invariants
(S0 : C0)	(P) $\mathtt{s}_S \sim \mathtt{Cstr}_{\mathtt{m}}^{\mathtt{char[]}}(\mathtt{s}_C)$
(S2 : C2)	(I1) $\mathtt{s}_S \sim \mathtt{Cstr}_{\mathtt{m}}^{\mathtt{char[]}}(\mathtt{s}_C)$
	(I2) $\mathtt{len}_S = \mathtt{i}_C$
(SE : CE)	(E) $\mathtt{ret}_S = \mathtt{ret}_C$

(b) Invariants Table for fig. 18a



(c) Product CFG for programs figs. 17a and 17c

PC-Pair	Invariants
(S0 : C0)	(P) $\mathtt{s}_S \sim \mathtt{Cstr}_{\mathtt{m}}^{\mathtt{clnode}}(\mathtt{cl}_C, 0)$
(S2 : C3)	(I1) $\mathtt{s}_S \sim \mathtt{Cstr}_{\mathtt{m}}^{\mathtt{clnode}}(\mathtt{cl}_C, 0)$
	(I2) $\mathtt{len}_S = \mathtt{i}_C$
(SE : CE)	(E) $\mathtt{ret}_S = \mathtt{ret}_C$

(d) Invariants Table fig. 18c

Figure 18: Product CFGs and Invariants Tables showing bisimulation between Strlen specification in fig. 17a and two C implementations in figs. 17b and 17c

5.1.3 Tree

We wrote a Spec program that sums all the nodes in a tree through an inorder traversal using recursion. We use two different data layouts for a tree: (1) a flat array where a complete binary tree is laid out in breadth-first search order commonly used for heaps ($\mathtt{Ctree}_{\mathtt{m}}^{\mathtt{u32[]}}$), and (2) a linked tree node with two pointers for the left and right children ($\mathtt{Ctree}_{\mathtt{m}}^{\mathtt{tnode(u32)}}$) (shown in table 6). Both Spec and C programs contain non-tail recursive procedure calls for left and right children. S2C is able to correlate these recursive calls using user-provided *Pre* and *Post*. At the entry of the recursive calls, S2C is required to prove that *Pre* holds for the arguments and at the exit of the recursive calls, S2C assumes *Post* on the returned states.

Table 7: Matrix and auxiliary List lifting constructors and their definitions.

Lifting Constructor	Definition
(T4) `Matrix = MNil | MCons(List, Matrix)`	
$\texttt{Cmat}_{\mathbb{m}}^{\texttt{u32[][]}}(p\ i\ u\ v{:}\texttt{i32})$	$\underline{\texttt{if}}\ i \geq_u u\ \ \underline{\texttt{then}}\ \texttt{MNil}$ $\underline{\texttt{else}}\ \texttt{MCons}(\texttt{Clist}_{\mathbb{m}}^{\texttt{u32[]}}(p[i]_{\mathbb{m}}^{\texttt{i32}}, 0_{\texttt{i32}}, v), \texttt{Cmat}_{\mathbb{m}}^{\texttt{u32[][]}}(p, i+1_{\texttt{i32}}, u, v))$
$\texttt{Clist}_{\mathbb{m}}^{\texttt{u32[r]}}(p\ i\ j\ u\ v{:}\texttt{i32})$	$\underline{\texttt{if}}\ j \geq_u v\ \ \underline{\texttt{then}}\ \texttt{LNil}$ $\underline{\texttt{else}}\ \texttt{LCons}(p[i \times v + j]_{\mathbb{m}}^{\texttt{i32}}, \texttt{Clist}_{\mathbb{m}}^{\texttt{u32[r]}}(p, i, j+1_{\texttt{i32}}, u, v))$
$\texttt{Cmat}_{\mathbb{m}}^{\texttt{u32[r]}}(p\ i\ u\ v{:}\texttt{i32})$	$\underline{\texttt{if}}\ i \geq_u u\ \ \underline{\texttt{then}}\ \texttt{MNil}$ $\underline{\texttt{else}}\ \texttt{MCons}(\texttt{Clist}_{\mathbb{m}}^{\texttt{u32[r]}}(p, i, 0_{\texttt{i32}}, u, v), \texttt{Cmat}_{\mathbb{m}}^{\texttt{u32[r]}}(p, i+1_{\texttt{i32}}, u, v))$
$\texttt{Clist}_{\mathbb{m}}^{\texttt{u32[c]}}(p\ i\ j\ u\ v{:}\texttt{i32})$	$\underline{\texttt{if}}\ j \geq_u v\ \ \underline{\texttt{then}}\ \texttt{LNil}$ $\underline{\texttt{else}}\ \texttt{LCons}(p[i + j \times u]_{\mathbb{m}}^{\texttt{i32}}, \texttt{Clist}_{\mathbb{m}}^{\texttt{u32[c]}}(p, i, j+1_{\texttt{i32}}, u, v))$
$\texttt{Cmat}_{\mathbb{m}}^{\texttt{u32[c]}}(p\ i\ u\ v{:}\texttt{i32})$	$\underline{\texttt{if}}\ i \geq_u u\ \ \underline{\texttt{then}}\ \texttt{MNil}$ $\underline{\texttt{else}}\ \texttt{MCons}(\texttt{Clist}_{\mathbb{m}}^{\texttt{u32[c]}}(p, i, 0_{\texttt{i32}}, u, v), \texttt{Cmat}_{\mathbb{m}}^{\texttt{u32[c]}}(p, i+1_{\texttt{i32}}, u, v))$
$\texttt{Cmat}_{\mathbb{m}}^{\texttt{lnode(u32[])}}(p\ v{:}\texttt{i32})$	$\underline{\texttt{if}}\ p = 0_{\texttt{i32}}\ \underline{\texttt{then}}\ \texttt{MNil}$ $\underline{\texttt{else}}\ \texttt{MCons}(\texttt{Clist}_{\mathbb{m}}^{\texttt{u32[]}}(p \xrightarrow{\mathbb{m}}_{\texttt{lnode}} \texttt{val}, 0_{\texttt{i32}}, v), \texttt{Cmat}_{\mathbb{m}}^{\texttt{lnode(u32[])}}(p \xrightarrow{\mathbb{m}}_{\texttt{lnode}} \texttt{next}, v))$
$\texttt{Cmat}_{\mathbb{m}}^{\texttt{lnode(u32)[]}}(p\ i\ u{:}\texttt{i32})$	$\underline{\texttt{if}}\ i \geq_u u\ \underline{\texttt{then}}\ \texttt{MNil}$ $\underline{\texttt{else}}\ \texttt{MCons}(\texttt{Clist}_{\mathbb{m}}^{\texttt{lnode(u32)}}(p[i]_{\mathbb{m}}^{\texttt{i32}}), \texttt{Cmat}_{\mathbb{m}}^{\texttt{lnode(u32)[]}}(p, i+1_{\texttt{i32}}, u))$
$\texttt{Cmat}_{\mathbb{m}}^{\texttt{clnode(u32)}}(p\ i\ u{:}\texttt{i32})$	$\underline{\texttt{if}}\ i \geq_u u\ \underline{\texttt{then}}\ \texttt{MNil}$ $\underline{\texttt{else}}\ \texttt{MCons}(\texttt{Clist}_{\mathbb{m}}^{\texttt{clnode(u32)}}(p[i]_{\mathbb{m}}^{\texttt{i32}}, 0_{\texttt{i2}}), \texttt{Cmat}_{\mathbb{m}}^{\texttt{clnode(u32)[]}}(p, i+1_{\texttt{i32}}, u))$

5.1.4 Matrix

We wrote a Spec program to count the frequency of a value appearing in a 2D matrix. A matrix is represented as an ADT that resembles a `List` of `List`s ((T4) in table 7). The C implementations for a `Matrix` object include (a) a two-dimensional array ($\texttt{Cmat}_{\mathbb{m}}^{\texttt{u32[][]}}$), (b) a flattened row-major array ($\texttt{Cmat}_{\mathbb{m}}^{\texttt{u32[r]}}$), (c) a flattened column-major array ($\texttt{Cmat}_{\mathbb{m}}^{\texttt{u32[c]}}$), (d) a linked list of 1D arrays ($\texttt{Cmat}_{\mathbb{m}}^{\texttt{lnode(u32[])}}$), (e) a 1D array of linked lists ($\texttt{Cmat}_{\mathbb{m}}^{\texttt{lnode(u32)[]}}$) and (f) a 1D array of chunked linked list ($\texttt{Cmat}_{\mathbb{m}}^{\texttt{clnode(u32)[]}}$) data layouts. Note that both `T[r]` and `T[c]` represent a 1D array of type `T`. The r and c simply emphasizes that these arrays are used to represent matrices in row-major and column-major encodings respectively. We also introduce two auxiliary lifting constructors, $\texttt{Clist}_{\mathbb{m}}^{\texttt{u32[r]}}$ and $\texttt{Clist}_{\mathbb{m}}^{\texttt{u32[c]}}$ for lifting each row of matrices lifted using the corresponding $\texttt{Cmat}_{\mathbb{m}}^{\texttt{u32[r]}}$ and $\texttt{Cmat}_{\mathbb{m}}^{\texttt{u32[c]}}$ `Matrix` lifting constructors. These constructors are listed in table 7.

Table 8: Equivalence checking times and minimum under- and over-approximation depth values at which equivalence checks succeeded.

Data Layout	Variant	Time(s)	($\mathbf{d}_u, \mathbf{d}_o$)	Data Layout	Variant	Time(s)	($\mathbf{d}_u, \mathbf{d}_o$)
	list				**tree**		
u32[]	sum naive	16	(1,2)	u32[]	sum	264	(1,2)
	sum opt	49	(4,5)	tnode(u32)	sum	204	(1,2)
	dot naive	65	(1,2)		**matfreq**		
	dot opt	176	(4,5)	u8[][]	naive	974	(1,3)
lnode(u32)	sum naive	8	(1,2)		opt	1.8k	(4,8)
	sum opt	54	(4,5)	u8[r]	naive	958	(1,3)
	dot naive	37	(1,2)		opt	1.9k	(4,8)
	dot opt	120	(4,5)	u8[c]	naive	984	(1,3)
	construct	426	(1,1)		opt	1.9k	(4,6)
clnode(u32)	sum opt	39	(4,5)	lnode(u8[])	naive	753	(1,3)
	dot opt	118	(4,5)		opt	1.7k	(4,6)
	strlen			lnode(u8)[]	naive	1.5k	(1,2)
u8[]	dietlibc$_s$	9	(1,2)		opt	2.3k	(4,6)
	dietlibc$_f$	44	(3,2)	clnode(u8)[]	opt	1.8k	(4,6)
	glibc	52	(3,2)		**strpbrk**		
	klibc	9	(1,2)	u8[],u8[]	dietlibc	398	(1,2)
	musl	49	(3,2)		opt	494	(4,2)
	netbsd	9	(1,2)	u8[],lnode(u8)	naive	392	(1,2)
	newlib	50	(3,2)		opt	540	(4,2)
	openbsd	8	(1,2)	u8[],clnode(u8)	opt	523	(4,2)
	uClibc	8	(1,2)	lnode(u8),u8[]	naive	497	(1,2)
lnode(u8)	naive	13	(1,2)		opt	602	(4,2)
	opt	49	(3,5)	lnode(u8),lnode(u8)	naive	345	(1,2)
clnode(u8)	opt	45	(3,5)		opt	503	(4,2)
	strchr			lnode(u8),clnode(u8)	opt	572	(4,2)
u8[]	dietlibc$_s$	16	(1,1)		**strcspn**		
	dietlibc$_f$	89	(4,1)	u8[],u8[]	dietlibc	462	(1,2)
	glibc	127	(4,1)		opt	538	(4,2)
	klibc	23	(1,1)	u8[],lnode(u8)	naive	395	(1,2)
	newlib$_s$	15	(1,1)		opt	521	(4,2)
	openbsd	24	(1,1)	u8[],clnode(u8)	opt	527	(4,2)
	uClibc	22	(1,1)	lnode(u8),u8[]	naive	601	(1,2)
lnode(u8)	naive	19	(1,1)		opt	660	(4,2)
	opt	146	(4,1)	lnode(u8),lnode(u8)	naive	349	(1,2)
	strcmp				opt	502	(4,2)
u8[],u8[]	dietlibc$_s$	39	(1,1)	lnode(u8),clnode(u8)	opt	595	(4,2)
	freebsd	39	(1,1)		**strspn**		
	glibc	41	(1,1)	u8[],u8[]	dietlibc	277	(1,2)
	klibc	41	(1,1)		opt	388	(4,2)
	musl	41	(1,1)	u8[],lnode(u8)	naive	405	(1,2)
	netbsd	39	(1,1)		opt	682	(4,2)
	newlib$_s$	42	(1,1)	u8[],clnode(u8)	opt	535	(4,2)
	newlib$_f$	405	(4,1)	lnode(u8),u8[]	naive	409	(1,2)
	openbsd	40	(1,1)		opt	553	(4,2)
	uClibc	38	(1,1)	lnode(u8),lnode(u8)	naive	357	(1,2)
lnode(u8),lnode(u8)	naive	47	(1,1)		opt	514	(4,2)
	opt	293	(4,1)	lnode(u8),clnode(u8)	opt	616	(4,2)
clnode(u8),clnode(u8)	opt	254	(4,1)				

5.2 Results

Table 8 lists the various C implementations and the time it took to compute equivalence with their specifications. For functions that take two or more data structures as arguments, we show results for different combinations of data layouts for each argument. We also show the minimum under-approximation (d_u) and over-approximation (d_o) depths at which the equivalence proof completed (keeping all other parameters to their default values).

During the verification of `strchr` and `strpbrk` implementations, we identified an interesting subtlety. Since `strchr` and `strpbrk` return null pointers to signify absence of the required character(s) in the input string, we additionally need to model the UB assumption that the zero address does not belong to the null character terminated array representing the string. We use an explicit constructor `SInvalid` to expose this well-formedness property in a Spec `String`. Furthermore, we relate `SInvalid` to the condition of C character pointer being null using the lifting constructors $\texttt{Cstr}^{\texttt{T}}_{\mathbb{m}}(p\texttt{:i32}, \dots)$ (as defined in table 5). These lifting constructors are used as part of Pre to equate S and C input strings. Finally in S, we model the absence of `SInvalid` in the input string as a UB assumption using the `assuming-do` statement introduced in section 4.1. Due to the (S `def`) assumption, this constraints the inputs to S as well as C to well-formed strings only. This is an example where (S `def`) and Pre can be used to model wellformedness of values in C.

TODO: add strlen spec atleast, show the strchr also!! maybe some matrix data layouts (only layouts)

6 Limitations

Our proof discharge algorithm is not without limitations. For a recursive relation relating values of a non-linear ADT such as `Tree`, a d-depth approximation results in $\sim 2^d$ smaller equalities. This is a major cause of inefficiency due to generation of large queries which slows down SMT solvers and counterexample-guided algorithms for large values of d.

S2C is only interested in finding a bisimulation relation and hence equivalence of non-bisimilar programs is beyond our scope. S2C currently only supports bitvector affine and inequality relations along with recursive relations provided as part of *Pre* and *Post*. Consequently, non-linear bitvector invariants (e.g. polynomial invariants) as well as custom recursive relations are not supported. While our correlation and invariant inference algorithms based on the Counter tool [22] are designed for translation validation between (C-like) unoptimized IR and assembly, we found them to be surprisingly good for Spec to (C-like) IR as well. Rather unsurprisingly, S2C suffers from the same limitations of these algorithms. For example, S2C supports path specializations from Spec to C, it does not search for path merging correlations.

7 Conclusion

As introduced in **??**, most of the current solutions to the problem of equivalence checking between a functional specification and a C program relies heavily on manually provided correlation, inductive invariants as well as proof assistants for discharging said obligations. While the size of programs considered in our work is quite small, we hope the ideas in S2C will help automate the proofs for such systems to some degree.

Prior work on push-button verification of specific systems [14, 35, 33, 34] involves a combination of careful system design and automatic verification tools like SMT solvers. Constrained Horn Clause (CHC) Solvers [17] encode verification conditions of programs containing loops and recursion, and raise the level of abstraction for automatic proofs. Comparatively, S2C further raises the level of abstraction for automatic verification from SMT queries and CHC queries to automatic discharge of proof obligations involving recursive relations.

A key idea in S2C is the conversion of proof obligations involving recursive relations to bisimulation checks. Thus, S2C performs *nested* bisimulation checks as part of a 'higher-level' bisimulation search. This approach of identifying recursive relations as invariants and using bisimulation to discharge the associated proof obligations may have applications beyond equivalence checking.

8 Outline of the Thesis

Chapter 1 of the thesis contains a general introduction to the research problem of verification C programs against a functional specification. We take a C program and its analogue in a safe functional language, and contrast their differences. We summarize our approach and finish with the major contributions.

Chapter 2 begins with an introduction to a minimal function language 'Spec' and an intermediate representation (IR). The rest of this chapter provides a background on bisimulation relation and product program, as well as introduce terminology used in the rest of the thesis. We finish with a formal definition of equivalence.

Chapter 3 starts with proof obligations and their properties. The rest of the chapter gradually introduces our first contribution: A Proof Discharge Algorithm and related sub-procedures with the help of two example programs introduced in the last two chapters. We also introduce a program representation of values, called 'deconstruction program'.

Chapter 4 contains a discussion on the two major components of our algorithm: (a) a counterexample-guided correlation algorithm to search for a bisimulation relation and (b) a counterexample-guided invariant inference algorithm. These two components along with our proof discharge algorithm allow automatic end-to-end equivalence checking. We formalize handling of procedure calls, and finish with a dataflow formulation of a pointer analysis used by our equivalence checker.

Chapter 5 introduces a program graph representation of values, called 'value graphs', similar to 'deconstruction program'. We motivate it by listing its advantages and give an algorithm to convert expressions to this representation. This helps us simplify our proof discharge algorithm.

In **Chapter 6**, we introduce our automatic equivalence checker tool named S2C, based on our proof discharge algorithm and counterexample-guided search procedures. S2C is evaluated on a large variety of C programs involving lists, strings, trees and matrices. This includes C programs taken from C library implementations as well as manually written programs. We show that our equivalence checker is able to prove equivalence of a single specification with multiple C implementations, each varying in its data layout and algorithmic strategy.

Finally, **Chapter 7** discusses the limitations of our algorithm and draws comparison with some related work. We note our key ideas and finish with potential improvements to our algorithm.

Publications Based on Research Work

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