

Exact Gap Computation for Code Coverage Metrics in ISO-C

Dirk Richter

Martin-Luther-University of Halle-Wittenberg, Germany
richter@informatik.uni-halle.de

Christian Berg

Martin-Luther-University of Halle-Wittenberg, Germany
christian.berg@student.uni-halle.de

Test generation and test data selection are difficult tasks for model based testing. Tests for a program can be meld to a test suite. A lot of research is done to quantify the quality and improve a test suite. Code coverage metrics estimate the quality of a test suite. This quality is fine, if the code coverage value is high or 100%. Unfortunately it might be impossible to achieve 100% code coverage because of dead code for example. There is a gap between the feasible and theoretical maximal possible code coverage value. Our review of the research indicates, none of current research is concerned with exact gap computation. This paper presents a framework to compute such gaps exactly in an ISO-C compatible semantic and similar languages. We describe an efficient approximation of the gap in all the other cases. Thus, a tester can decide if more tests might be able or necessary to achieve better coverage.

1 Introduction

Tests are used in model based testing to identify software defects. High quality test generation and test data selection can be difficult tasks when the test has to satisfy a lot of requirements or cannot be created automatically because of the undecidability of the halting problem in Turing powerful languages. Given requirements for a test suite (set of tests) are functional or non-functional (e.g. execution times, runtime, usage of memory, correctness or a minimum value of a code coverage metric). Code coverage metrics quantify the quality of a test suite rather imprecisely and guide testers only. There is a gap between the feasible and theoretical maximal possible code coverage value. Sometimes demanded requirements are unsatisfiable because of gaps. Unnecessary additional tests will be computed while not all requirements are satisfied. This enlarges the test suite and introduces redundancy. Fortunately these problems (caused by metric imprecisions) can be solved by computing these gaps, which is not possible for Turing powerful languages in general. Therefore this paper presents suitable models in a new C-like syntax. These models allow to use an ISO-C compatible semantic. In this paper we show how to compute such gaps exactly for these models using formal verification techniques resp. software model checking ideas. The paper is organized as follows: at first we clarify basics and used notations; we then present our framework, apply it to some common coverage metrics and illustrate this on some examples. Finally we discuss related work and present a summary and conclusions.

2 Basics

2.1 Code Coverage Metrics γ

Let $T_P = 2^{tests}$ be the set of all possible sets of tests and P a program written in a common programming language such as C, C++ or Java. Each $t = \{\alpha_1, \alpha_2, \dots\} \in T_P$ is a test suite with tests α_i for program P . The function $\gamma^P : T_P \rightarrow [0, 1]$ is a code coverage metric, if γ^P is monotonically increasing. The program P can be omitted, if it is well-defined by the context.

In this paper some common code coverage metrics for functions, statements, decisions, branches and conditions will be considered as examples. Other ones (e.g. linear code sequence and jump coverage, jj-path coverage, path coverage, entry/exit coverage or loop coverage) can be adapted in a similar way.

The function coverage metric $\gamma_f^P(t) := |func(t)|/|func(P)|$ is the ratio of functions $func(t)$ that has been called in the test suite t , to all functions $func(P)$ in P [12].

The statement coverage metric $\gamma_s^P(t) := |stats(t)|/|stats(P)|$ is the ratio of statements $stats(t)$ that has been executed in the test suite t , to all statements $stats(P)$ in P [12]. To distinguish the same statement s on different program points l_1 and l_2 , we annotate each statement s with unique labels l_1 and l_2 from program P , so that $l_1 : s \in stats(P)$ and $l_2 : s \in stats(P)$.

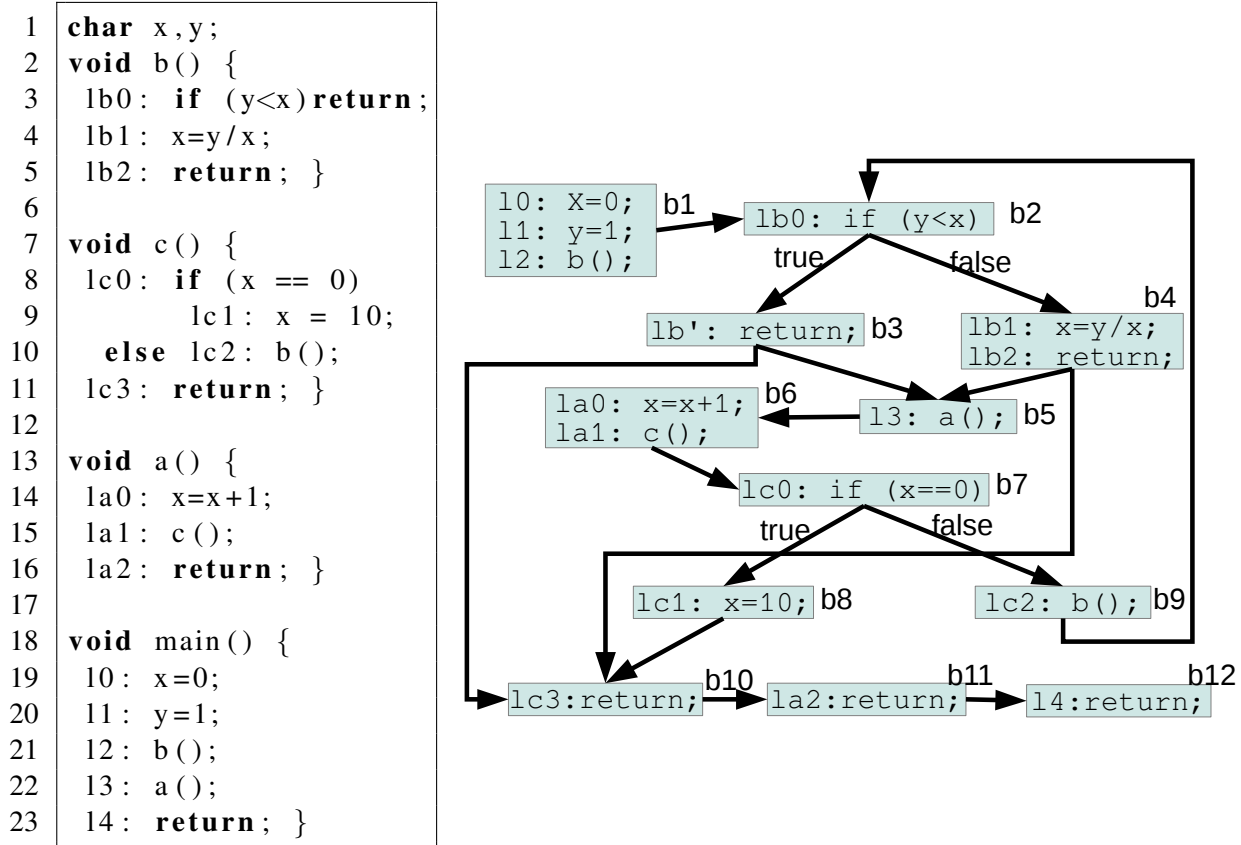


Figure 1: SPDS example P_1 in ISO-C syntax and corresponding BBICFG

Let $blocks(P)$ be all basic blocks [1] in P . Every program point has a surrounding basic block. The basic block inter-procedural control flow graph $BBICFG_P = (blocks(P), edges(P))$ (see Fig. 1) consists of the basic blocks $blocks(P)$ as nodes and edges $edges(P) \subseteq blocks(P)^2$, where $(b_1, b_2) \in edges(P)$ iff there is an execution path of length 1 from the end of block b_1 to the entry of block b_2 (execution of the last statement of block b_1). The decision coverage metric $\gamma_d^P(t) := |edges(t)|/|edges(P)|$ is the ratio of executed edges $edges(t)$ of the control flow graph $BBICFG_P$ for t , to all edges $edges(P)$ in P .

The branch coverage metric $\gamma_b^P(t) := |blocks(t)|/|blocks(P)|$ is the ratio of basic code blocks $blocks(t)$ executed during test suite t , to all basic blocks $blocks(P)$ in P [13]. Even if all basic blocks are covered by test suite t and $\gamma_b^P(t) = 1$, there can be uncovered branching edges in the basic block inter-procedural control flow graph $BBICFG_P$. Thus $\gamma_d^P(t) < 1$ is possible in this case.

Let $bExpr(l)$ be the set of all Boolean sub-expressions on label l of program P and

$$BExpr(P) := \{(l, bExpr(l)) \bullet l \in labels(P)\}. \quad (1)$$

The condition or predicate coverage metric $\gamma_c^P(t) := |exval(t, P)| / (2 \cdot |BExpr(P)|)$ is the ratio of evaluations of boolean sub-expressions $exval(t, P) \subseteq BExpr(P) \times \{true, false\}$ of the test suite t , to all evaluations of boolean sub-expressions in P [12]. The relation $exval(t, P)$ describes the evaluations of sub-expressions e on label l under test suite t , such that $((l, e), true) \in exval(t, P)$ iff there is a test $\alpha \in t$ where e can be evaluated to $true$ on label l under test α . When boolean operations are not short circuited, condition coverage does not necessarily imply decision coverage.

2.2 Code Coverage Metric Gap δ

Let $\gamma: T_P \rightarrow [0, 1]$ be a code coverage metric for a program P . The code coverage metric gap $\delta_\gamma(P) \in [0, 1]$ is the smallest difference between the coverage ratio of a test suite $t \in T_P$ and the theoretical maximal value 1:

$$\delta_\gamma(P) := \inf_{t \in T_P} (1 - \gamma(t)). \quad (2)$$

Let $\delta_x(P)$ denote $\delta_{\gamma_x}(P)$, where $x \in \{c, d, s, f, b\}$. Obviously dead code can cause $\delta_\gamma(P) > 0$. If some evaluations of boolean (sub-)expressions cannot be realized, $\delta_\gamma(P) > 0$ is possible without dead code (e.g. condition or decision coverage). In Turing powerful programming languages the gap $\delta_\gamma(P)$ can not be computed in general, because the halting problem is undecidable. In this case the gap $\delta_\gamma(P)$ can be approximated only. We show how to compute the exact gap $\delta_\gamma(P)$ for an ISO-C compatible semantic by adequate modeling.

2.3 Suitable Models

A more expressive model describing the program behaviour allows for a more accurate approximation of the gap $\delta_\gamma(P)$. If the model is not Turing powerful and the model behaviour is equivalent to the program behaviour, the gap $\delta_\gamma(P)$ can be exactly computed. Therefore we defined an ISO-C compatible semantic using pushdown systems (PDS) [9]. The split of the ISO-C language definition into platform-independent semantics and platform-specific semantics has a serious implication for deciding the halting problem of ISO-C programs: whether a C program halts or not depends on the platform-specific semantic. Thus, even though the halting problem for a C program is decidable for a platform-specific semantic, the halting property can become undefined if no specific platform is assumed [9]. Now we present an extension of the PDS used in [9] to symbolic pushdown systems (SPDS) using an ISO-C like syntax. SPDS use a more compact representation and define the PDS configurations and transitions symbolically.

A SPDS is a tuple $S = (vgbl, func)$, where $vgbl$ is a finite set of variables (global variables in ISO-C) and $func$ is a set of functions (pairwise different names) with an initial function $main \in func$. Each variable v has an integer type¹ $bits(v) \in \mathbb{N}_{\geq 1}$ and a fixed length $len(v) \in \mathbb{N}_{\geq 1}$. Every variable v is an array. A function is a tuple $(f, param, vlcl, stats)$, where $(f, param)$ is a function signature with a unique function name f and a finite list of parameter variables $param$. The set $vlcl$ is a finite set of variables (local variables in ISO-C), such that $param \subseteq vlcl$. The body of f is a finite list of statements $stats$. Each statement $l : s \in stats$ has a unique label $l \in labels(f)$, and $fst(f) \in labels(f)$ is the label of the first statement in the list $stats$. We use $vgbl = vgbl(S), func = func(S), param = param(f), vlcl = vlcl(f)$

¹The Boolean constants *false* and *true* are represented via 0 and $\neq 0$ like ISO-C.

and $stats = stats(f)$ respectively, if S or f are well-defined by the context. Let denote $func(l)$ the function f for which $l \in labels(f)$. Further let be

$$vars := vgbl \cup \bigcup_{f \in func} vlcl(f), \quad stats(S) := \bigcup_{f \in func} stats(f) \quad \text{and} \quad labels(S) := \bigcup_{f \in func} labels(f). \quad (3)$$

Similar to ISO-C the SPDS variables are used to build expression $Expr$ using constants and operators. The priority and associativity are the same as in ISO-C.

An expression $e \in Expr$ can be strictly evaluated to an integer number $\llbracket e \rrbracket_g^{c_f} \in \mathbb{Z} \cup \{\perp\}$ using valuation functions for global and local variables $g : vgbl \times \mathbb{Z} \rightarrow \mathbb{Z}$ and $c_f : vlcl(f) \times \mathbb{Z} \rightarrow \mathbb{Z}$. The symbol \perp denotes arithmetic exception (e.g. division-by-zero or index-out-of-bounds). The functions $g(v, i)$ and $c_f(v, i)$ return the current value of variable v at index i (value of $v[i]$). The evaluation functions g and c_f can be omitted, if they are well-defined by the context. A variable usage $v[i]$ of variable $v \in vars$ with index $i \in \mathbb{Z}$ is evaluated as

$$\llbracket v[i] \rrbracket_g^{c_f} := \begin{cases} g(v, i) & v \in vgbl \wedge 0 \leq i < len(v) \\ c_f(v, i) & v \in vlcl(f) \wedge 0 \leq i < len(v) \\ \perp & otherwise. \end{cases} \quad (4)$$

For $e \in Expr$ and a statement $l : s \in stats(f)$, s has one of the following forms:

- $v[e_1] = e_2$; corresponds to writing the value $\llbracket e_2 \rrbracket$ into the variable v at index $\llbracket e_1 \rrbracket$.
- $f(v_1, \dots, v_n)$; corresponds to a function call (call by value), iff $(f, param)$ is a signature, where $param = [p_1, p_2, \dots, p_n]$, $v_i \in vars$ and $bits(v_i) \leq bits(p_i)$ for all $1 \leq i \leq n$.
- *return*; corresponds to a function return.
- *if (e) goto l'*; corresponds to a conditional jump² to label $l' \in labels(f)$.

The exception of a dynamic type mismatch occurs for " $v[e_1] = e_2$;" and the system terminates, if $\llbracket e_2 \rrbracket = \perp$ or the type $bits(v)$ is too small to store the value $\llbracket e_2 \rrbracket$ or $\llbracket e_1 \rrbracket \notin \{0, 1, \dots, len(v)\}$. We denote $v = e$ for $v[0] = e$ and v for usages of $v[0]$ to emulate syntactically non-array variables. The system terminates on statement "*if (e) goto l'*;" too, if $\llbracket e \rrbracket = \perp$. The predefined function $rand(e)$ returns a random number between 0 and $\llbracket e \rrbracket$ for $e \in Expr$, whereby $rand(\perp) = \perp$. Further ISO-C statements and variations for other languages can be mapped to these basic statements in the modeling phase. All variables (global and local) are uninitialized and have initially a random value. A test α for S is a subset of global variables with predefined values for label $fst(main)$. A configuration $s = (g, [(l_n, c_n), (l_{n-1}, c_{n-1}), \dots, (l_1, c_1)])$ of S represents a state of the underlying Kripke structure with the current execution label $l_n \in labels$, the valuation $g : vgbl \times \mathbb{Z} \rightarrow \mathbb{Z}$ of global variables and the stack content. The stack content consist of a list of function calls with current execution labels $l_i \in labels(S)$ and valuations for local variables $c_i : vlcl(func(l_i)) \times \mathbb{Z} \rightarrow \mathbb{Z}$. The head of s is $head(s) = (g, (l_n, c_n))$. The set of all possible configurations is $conf(S)$. A *run* of S is a sequence of consecutive configurations beginning with an initial configuration $(g_{init}, [fst(main), c_{init}]) \in conf(S)$. SPDS are (like PDS) not Turing powerful and can be used to model the behaviour of (embedded) ISO-C programs. There is no restriction on the recursion depth.

²intra-procedural

3 Exact Gap Computation Framework

Let $\gamma: T_P \rightarrow [0, 1]$ be a code coverage metric for a program P . Our framework to compute the gap $\delta_\gamma(P)$ consists of the following steps:

1. If necessary, create a SPDS model S with ISO-C compatible semantic for program P .
2. Modify the model S to a SPDS model S' to enable gap analysis for the code coverage metric γ .
3. Compute exact variable ranges for some new variables in S' .
4. Conclude the exact size of the gap $\delta_\gamma(S)$ in S for the code coverage metric γ .
5. Conclude the size of the gap $\delta_\gamma(P)$ in P .

3.1 SPDS Modeling (step 1)

If the given program P is not written in ISO-C (e.g. Java) or P has another platform-specific semantic, we create a SPDS model S for P by abstraction. Otherwise the behaviour of S is the same of P by mapping all the ISO-C statements to the basic SPDS-statements of section 2.3 using abbreviations (described in this section). Java can be handled using the tool JMoped [7]. Often other languages and corresponding statements can be mapped to the basic SPDS-statements in a similar fashion. For simplification we present some common mappings, which are abbreviations for previously defined basic SPDS-statements. We sketch the ideas only, because of limited space. Fig. 1 shows the SPDS example P_1 in ISO-C syntax, where "char x" in line 1 is an abbreviation for "int x(8) [1]" to declare an integer array of type $bits(x) = 8$ and $len(x) = 1$.

Omitted Returns and Labels: If there is no return statement at the end of a function body, its existence is assumed during interpretation of the symbolical description of S . The same holds for statements without labels, such that each statement in the SPDS has a unique label after interpretation.

Parameter Expressions: Basic SPDS-statements allow variables to be passed as parameters in function calls. We can simulate to pass expressions by temporary local variables. Let " $l: f(e_1, e_2, \dots, e_n)$;" be a function call with expressions $e_i \in Expr$, where $(f, param)$ is a signature with $param = [p_1, p_2, \dots, p_n]$. We introduce new local SPDS-Variables $pe_i \notin vars$ with type $bits(pe_i) = bits(p_i)$ and $len(pe_i) = 1$. These variables are used to evaluate the expressions before the function call: $pe_i = e_i$. Instead of e_i now pe_i is passed to f using the basic SPDS-statement $f(pe_1, pe_2, \dots, pe_n)$. The function call " $l: f(e_1, e_2, \dots, e_n)$ " is interpreted as " $l: pe_1 = e_1; pe_2 = e_2; \dots; pe_n = e_n; f(pe_1, pe_2, \dots, pe_n)$ ". Now only basic SPDS-statements are used. The code coverage metrics are adapted accordingly. For example the statements $pe_i = e_i$; are ignored for the statement coverage metric.

Return Values: A function can return a value. This value can be used to set a variable " $v = f(\dots)$ ";. If a function returns an expression e via " $return e$ ", a new global variable $ret_f \notin vars$ is introduced. The type of ret_f equals the return type of function f and $len(ret_f) = 1$. The statement " $return e$;" is interpreted as " $ret_f = e; return$ ";. On the other hand the assignment " $v = f(\dots)$;" is interpreted as " $f(\dots); v = ret_f$;" to store the return value of f in v .

Function Calls in Expressions: If there is a function call $f(\dots)$ in an expression e , this function is evaluated in a temporary local variable. Boolean operations in SPDS are strict and not short circuited. Short circuited expressions (also increments $i++$ and decrements $i--$) can be mapped to strict expressions without side effects by several conditional statements. Thus every function call $f(\dots)$ in an expression e will be definitively evaluated during the evaluation of e . Accordingly it is safe to do every call before the evaluation of e . Sometimes the order of this evaluation is implementation defined (as in ISO-C) and

depends on the source language (e.g. `i++*i++`). Thus we use priority and associativity for calculating this order. The intermediate representation of a compiler can be used, too, to achieve a mapping to SPDS.

Unconditional Jump: The statement `goto l;` is mapped to `if (1) goto l;`. The dead branching edge to the following statement is ignored by the decision coverage etc.

Skip Statement: Particularly low level languages have often a "no operation" statement. We use the statement `skip`, which does not change variable settings. The statement `l : skip;` can be interpreted using the conditional branch `l : if (0) goto l;`. If there is a global variable $v \in vars$, this can also be interpreted as `l : v = v;`. The former needs no consideration for the coverage metrics.

Random Numbers: In ISO-C the function `rand()` returns a pseudo random value between 0 and the constant `RAND_MAX`, where `RAND_MAX` depends on the system. We can map this behaviour using the SPDS function `rand(RAND_MAX)`. A similar mapping for random numbers is possible in other languages.

Conditional Statements: Let $s1$ and $s2$ be lists of statements. The conditional statement `if (e) s1 else s2;` is interpreted as `if (e) goto l1;s2;goto l';l1 : s1;l' : skip;`, where $l_1, l' \notin labels$. `if (e) s;` is an abbreviation for `if (e) s else skip;`.

Local Variable Definitions: A local variable can be defined during an assignment of a basic block or a loop header. Such local variable definitions are mapped to local variables of the surrounding function. Renaming can be done easily if necessary.

Loops: A for-loop of the form `for (init; cond; inc) body;` is interpreted as `init; l : body; inc; if (cond) goto l;`, where $l \notin labels$. The `do` and `while` loops are interpreted in a similar way.

Modular Arithmetic and Integer Overflow: The ISO-C standard says that an integer overflow causes "undefined behaviour", meaning that compilers conforming to the standard can generate any code: from completely ignoring the overflow to aborting the program. Our solution terminating the system is conform to the ISO-C standard. Evaluations of expressions in SPDS are not restricted to arithmetic bounds, but dynamic type mismatches are possible for assignments $v = e$. In the case of modeling nonterminating modular arithmetic the modulo operator `%` can be used to shrink the expression e to fit the size $bits(v)$. Hence, a dynamic type mismatch does not occur.

Dynamic Memory and Pointers: In ISO-C a certain amount of the heap can be reserved using the function `malloc(int)`. It returns an address on the heap. The heap is finite, because the number of addresses is finite. This behaviour is simulated using a global array `heap` of type $bits(heap) = 8$ with length $len(heap) = m$ and a global variable `ptr` with type $bits(ptr) = \lceil \log_2(m) \rceil$ and $len(ptr) = 1$, which points to the next free space in the heap array. The function `malloc(int)` can be implemented as shown in Listing 1 with 1024 heap elements respectively, which needs a 10 bit variable `ptr` for accessing. A memory exceptions occurs (label `memout`), if there is not enough memory left.

```

1  int heap(8)[1024];
2  int ptr(10);
3
4  int(10) malloc(int n(10)) {
5      if (ptr >= 1024-n) goto memout;
6      ptr = ptr+n;
7      return ptr-n; }

```

Listing 1: Malloc as SPDS in ISO-C like syntax

Once reserved space can be reused, because a garbage collector and a function `free` can be implemented in SPDS. A pointer is a SPDS variable used as an index of the heap array and an address is just another index (returned by the address operator `&`). Variables placed in the heap array support the address

operator in contrast to the other SPDS variables. If putting a **local** variable of a function f into the heap array, the recursion of f will be bound, because of a finite maximal heap size. Coverage metrics have to adapt to these additional SPDS functions, statements and variables to be able to compute the correct gap. **Call by Reference:** Instead of passing a variable as a function parameter, a pointer can be used to indirectly access variable values in the heap. Thus call by reference can be simulated. Unfortunately this results in bounding the recursion, too.

Dynamic Arrays: Array semantic in ISO-C is defined by pointers and access to its elements is defined by pointer arithmetic. Thus $malloc(int)$ can be used for this purpose.

Other constructs and statements from other languages (e.g. classes, structs, objects, dynamic parameter lists, etc.) can be mapped in a similar way. If an arithmetic exception occurs, the SPDS ends and the corresponding ISO-C program P can have undefined behaviour according to the language specification. P can terminate, which is a complying behaviour. Therefore this behaviour is used for our modeling process. Other implementation defined behaviour can be modeled similarly.

3.2 Extraction of Exact Variable Ranges (step 3)

In step 2 a SPDS S' is created for the SPDS S by slightly modifying S (explained in the next section). For a PDS B an automaton $Post^*(B)$ can be computed, which accepts all reachable configurations of B [18]. Thus for the SPDS S' a similar automaton $Post^*(S')$ can be created, because S' is just a symbolical PDS. This is a basic step in symbolic model checking using Moped [14]. We use the $Post^*$ algorithm of the model checker Moped for our implementation by mapping our SPDS definitions to the input language Remopla³ [17]. The set of reachable heads $h(S') := \{(g, (l, c)) \bullet (g, [(l, c)...]) \in Post^*(S')\}$ is finite because of finite variable types. Thus exact variable ranges can be extracted from $h(S')$. Let $v \in vars$ and $l \in labels$, then $range_l^{S'}(v) := \{[[v]]_g^c \bullet (g, (l, c)) \in h(S')\}$ is the exact variable range of v . The notation S' can be omitted, if S' is well-defined by the context. For all values $k \in range_l(v)$ there is a run of S' , such that $[[v]] = k$ on label l and vice versa.

$h(S')$ and $range_l(v)$ can be computed symbolically out of $Post^*(S')$ using Ordered Binary Decision Diagrams (OBDD) operations. The computation of $h(S')$ is a straightforward OBDD restriction operation in $Post^*(S')$ and results in a characteristic function $q : \{0, 1\}^n \rightarrow \{0, 1\}$ represented as an OBDD. The input vectors of q are heads $h(S')$ encoded as finite Bit sequences. The computation of $range_l(v)$ uses cofactors. A cofactor of q is $q[x_i = b](x_1, x_2, \dots, x_n) := q(x_1, x_2, \dots, x_{i-1}, b, x_{i+1}, \dots, x_n)$ [8]. The positive cofactor is $q[x_i = 1]$ and the negative cofactor is $q[x_i = 0]$. A characteristic function $r : \{0, 1\}^m \rightarrow \{0, 1\}$ for $range_l(v)$ can be computed using cofactors:

Lemma 3.1 *Let k be the starting index of the encoding of v on label l in q and let m be the length of the encoding. Then $r(y_1, y_2, \dots, y_m) = 1$ is valid, iff $q[x_k = y_1][x_{k+1} = y_2] \dots [x_{k+m-1} = y_m]$ is not always 0 (not the empty OBDD).*

The proof is a consequence of the definitions. The computation of exact variable ranges is more time-consuming than model-checking the reachability in S' [8]. Fortunately $range_l(v)$ can be approximated using static data flow analyses and test suites. This is the case for focusing on efficiency or unbounded recursion depth in combination with unbounded parallelism. For further reference in comparisons, explanations, and proofs see [8].

3.3 SPDS Supplementation (step 2) and Exact Gap Inference (step 4)

Now we show exemplary, how to apply our framework to common code coverage metrics.

³e.g. we map integers to nonnegative numbers, as Remopla does not support negative integers

3.3.1 Function Coverage Gap $\delta_f(S)$

We supplement S with a new global variable $v \notin \text{vgbl}(S)$ using type $\text{bits}(v) = 1$ and $\text{len}(v) = 1$ without any assignment or reading usage on v to ensure the existence of at least one global variable in S' . By construction this variable v has a random undefined value $\llbracket v \rrbracket \in \{0, 1\}$ on each label $l \in \text{label}(S')$ resp. on each program point. The exact function coverage gap $\delta_f(S)$ can be concluded from the exact ranges of variables in S' as follows:

Lemma 3.2

$$\delta_f(S) = 1 - \frac{|\{f \in \text{func}(S) \bullet \text{range}_{fst(f)}^{S'}(v) \neq \emptyset\}|}{|\text{func}(S)|} \quad (5)$$

Proof (sketch): The new global variable v does not influence the model behaviour. All variable evaluations and reachable labels in S' are the same in S . Further it is $\text{vgbl}(S') = \text{vgbl}(S) \cup \{v\}$, $\text{func}(S') = \text{func}(S)$ and $\text{func}(S) \neq \emptyset$ because of $\text{main} \in \text{func}(S)$. The main observation is, that a label $l \in \text{labels}(S)$ is unreachable or dead, iff $\text{range}_l^{S'}(v) = \emptyset$. Thus a function f can be called, iff $\text{range}_{fst(f)}^{S'}(v) \neq \emptyset$. Choose a test suite $t' \in T_S$ such that $|\text{func}(t')|$ is maximal. With the maximality of t' we have

$$|\text{func}(t')| \geq |\{f \in \text{func}(S) \bullet \text{range}_{fst(f)}^{S'}(v) \neq \emptyset\}|. \quad (6)$$

Every function $f \in \text{func}(t')$ has a cover witness test $\alpha \in t'$, so that the label $\text{fst}(f)$ is reachable under test α . Thus it is $\text{range}_{fst(f)}^{S'}(v) \neq \emptyset$. On the other hand we obtain

$$|\text{func}(t')| \leq |\{f \in \text{func}(S) \bullet \text{range}_{fst(f)}^{S'}(v) \neq \emptyset\}|, \quad (7)$$

because each $f \in \text{func}(S)$ with $\text{range}_{fst(f)}^{S'}(v) \neq \emptyset$ has at least one test α' (not necessarily $\in t'$) to cover the function f , which can be detected by an evaluation of v . Accordingly it is

$$\sup_{t \in T_S} |\text{func}(t)| = |\{f \in \text{func}(S) \bullet \text{range}_{fst(f)}^{S'}(v) \neq \emptyset\}|, \quad (8)$$

which is equivalent to

$$\inf_{t \in T_S} (1 - \gamma_f(t)) = 1 - \frac{|\{f \in \text{func}(S) \bullet \text{range}_{fst(f)}^{S'}(v) \neq \emptyset\}|}{|\text{func}(S)|}. \quad (9)$$

□

The exact branch coverage gap $\delta_b(S)$ can be computed similarly. Instead of function entry points, just the block entry points are considered.

3.3.2 Statement Coverage Gap $\delta_s(S)$

The SPDS S' will be supplemented with a new variable $v \notin \text{vgbl}(S)$ and the type $\text{bits}(v) = 1$ and $\text{len}(v) = 1$ similarly to the function coverage gap. The exact statement coverage gap $\delta_s(S)$ can be computed:

Lemma 3.3

$$\delta_s(S) = 1 - \frac{|\{l : s \in \text{stats}(S) \bullet \text{range}_l^{S'}(v) \neq \emptyset\}|}{|\text{stats}(S)|} \quad (10)$$

Proof (sketch): Similar to Lemma 3.2 prove $\sup_{t \in T_S} |\text{stats}(t)| = |\{l : s \in \text{stats}(S) \bullet \text{range}_l^{S'}(v) \neq \emptyset\}|$.

□

3.3.3 Decision Coverage Gap $\delta_d(S)$

The branch coverage uses the nodes of the control flow graph $BBICFG_S$ and the decision coverage uses the edges. The execution of an edge $(b_1, b_2) \in edges(S)$ in the control flow graph $BBICFG_P$ depends on several conditions such as arithmetic overflow, division-by-zero or boolean expressions for conditional branches. To compute the exact decision coverage gap, we introduce a new global variable $v_{in} \notin vars(S)$ into S' with type $bits(v_{in}) = 1 + \lceil \log_2(|blocks(S)|) \rceil$ and $len(v_{in}) = 1$. Each label l belongs to a basic block b_l , which can be identified by a unique number $n_{b_l} \in \mathbb{N}_{\geq 0}$. This number is assigned to the variable v_{in} to detect the past basic block for a statement. The type $bits(v_{in})$ is big enough to store every unique identifier n_{b_l} . Each statement " $l : s$ " $\in stats(S)$ is modified⁴ to " $l : v_{in} = n_{b_l}; l' : s$ " in S' , where $l' \notin labels(S)$ is unique. So it is possible to determine the past basic block on label l using the exact range of the SPDS variable $v_{in} \in vgl(S')$. The exact decision coverage gap $\delta_d(S)$ can be computed:

Lemma 3.4

$$\delta_d(S) = 1 - \frac{|\{(a, b) \in edges(S) \bullet n_a \in range_{fst(b)}^{S'}(v_{in})\}|}{|edges(S)|} \quad (11)$$

Proof (sketch): Let $a, b \in blocks(S)$ be basic blocks. By construction it is $n_a \in range_{fst(b)}^{S'}(v_{in})$, iff there is an execution path from the end of basic block a directly to the first label $fst(b)$ of basic block b in S . This is equivalent to the existence of a test α , such that $(a, b) \in edges(\alpha)$. Thus we have

$$\exists \alpha \in t' : (a, b) \in edges(\alpha) \Leftrightarrow n_a \in range_{fst(b)}^{S'}(v_{in}) \quad (12)$$

for a chosen $t' \in T_S$, where $|edges(t')|$ is maximal. Hence it is

$$\sup_{t \in T_S} |edges(t)| = |\{(a, b) \in edges(S) \bullet n_a \in range_{fst(b)}^{S'}(v_{in})\}|, \quad (13)$$

which shows (11) similar to Lemma 3.2. □

3.3.4 Condition Coverage Gap $\delta_c(S)$

For the condition coverage all boolean sub-expressions (conditions $BExpr(S)$) on each label are considered. The theoretical maximal value can be achieved, when every condition $(l, e) \in BExpr(S)$ can be 1 (*true*) and 0 (*false*). For each boolean sub-expression $b \in B = \bigcup_{l \in labels(S)} bExpr(l)$ we introduce new boolean global variables $v_b \notin vars(S)$ with type $bits(v_b) = 1$ and $len(v_b) = 1$ into S' . Let further $bExpr(l) = \{e_1, e_2, \dots, e_n\}$ be the set of all boolean sub-expressions on label l . Each statement " $l : s$ " $\in stats(S)$ with $bExpr(l) \neq \emptyset$ will be modified to " $l : v_{e_1} = e_1; v_{e_2} = e_2; \dots, v_{e_n} = e_n; l' : s$ " in S' , where $l' \notin labels(S)$ is unique. The statement " $l : s$ " $\in stats(S)$ will be modified to " $l : skip; l' : s$ " in S' , if $bExpr(l) = \emptyset$. Hence the existence of label $l' \in labels(S')$ is guaranteed. Thus the exact condition coverage gap $\delta_c(S)$ can be computed:

Lemma 3.5

$$\delta_c(S) = 1 - \frac{\sum_{\substack{l \in labels(S) \\ e \in bExpr(l)}} |range_{l'}^{S'}(v_e)|}{2 \cdot |BExpr(S)|} \quad (14)$$

⁴Additionally this can be done using the native synchronous parallelism in SPDS without an extra label: " $l : v_{in} = n_{b_l}, s$ ".

Proof (sketch): Choose a $t' \in T_S$ such that $|exval(t', S)|$ is maximal. Then it is

$$((l, e), b) \in exval(t', S) \quad (15)$$

$$\Leftrightarrow \text{expression } e \text{ can be evaluated to } b \in \{0, 1\} \text{ on label } l \text{ in } S \quad (16)$$

$$\Leftrightarrow b \in range_{l'}^S(v_e). \quad (17)$$

This proves (14) similar to Lemma 3.2, because of

$$\sup_{t \in T_S} |exval(t, S)| = \sum_{\substack{l \in labels(S) \\ e \in bExpr(l)}} |range_{l'}^S(v_e)|. \quad (18)$$

□

3.4 Conclusion for $\delta_\gamma(P)$ based on $\delta_\gamma(S)$ (step 5)

The computed gap $\delta_\gamma(S)$ is exact ($\delta_\gamma(S) = \delta_\gamma(P)$), if the behaviour of S is equivalent to the behaviour of P . Hence step 1 does not abstract nor simplify the program behaviour. This is the case for our ISO-C compatible semantic [9] on C programs. If S abstracts from the behaviour of P , $\delta_\gamma(S)$ is only an approximation for $\delta_\gamma(P)$. The approximation degree depends on the degree of this abstraction.

4 Gap Approximation using δ_γ^- and δ_γ^+

The gap can be approximated by abstracting the program P to a simpler behavior of SPDS S as shown above. On the other hand the exact variable ranges $range_l(v)$ can be approximated, too. This is a more practical approach particularly for huge software systems. Let $range_l^+(v)$ be an over- and $range_l^-(v)$ an under-approximation of $range_l(v)$. The sets $range_l^-(v)$ can be realized using a test suite $t \in T_P$. All occurring variable values during the tests $\alpha \in t$ can be used as lower bound for $range_l(v)$. On the other hand $range_l^+(v)$ can be realized using a conservative data flow analysis. This usually results in additional variable values, which never can be achieved. Both δ_γ^- and δ_γ^+ can be defined similar to δ_γ using $range_l^-(v)$ and $range_l^+(v)$ instead of $range_l(v)$. It is easy to realize, how to bound the gap δ_γ using $range_l^-(v)$ and $range_l^+(v)$:

Lemma 4.1 $\delta_\gamma^+ \leq \delta_\gamma \leq \delta_\gamma^-$.

Obviously the gap approximation is perfect and an exact gap is found, if $\delta_\gamma^+ = \delta_\gamma^-$. In this case it is not necessary to compute exact variable ranges.

5 Exemplary Illustration

As a comparative measurement of our method the values calculated by gcov [22]⁵ are presented at the end of this section. The free tool gcov calculates the code coverage during an execution, which can be used to track the code coverage of a test suite.

To show the concepts presented so far, we use example P_1 in Fig. 1 and example P_2 in Fig 2. The constructs in the presented ISO-C code are automatically mapped to SPDS-statements as described in section 3.1. P_1 contains an arithmetic exception, caused by a division-by-zero. Hence P_1 contains a lot of dead code and any test suite with at least one test would be complete (i.e. there is no way to

⁵<http://gnu.org/software/gcov>

cover more code). There is no test necessary ($t = \emptyset$), because the variables x and y are initialized on labels $l0$ and $l1$. Thus it is $range_l(x) = range_l(x)^- = \{0\}$ and $range_l(y) = range_l(y)^- = \{1\}$ for all $l \in L$, where $L = \{l0, l1, l2, lb0, lb1\}$. It is $range_l(x) = range_l(x)^- = range_l(y) = range_l(y)^- = \emptyset$ for all $l \in labels(P_1) \setminus L$. All the conditions in conditional branches are considered to be statements (see BBICFG in the right part of Fig. 1), because a conditional branch can contain a statement (e.g. $x=0$ in "if ($x=0$)...").

Thus it is $|stats(P_1)| = 15$, $|blocks(P_1)| = 12$, $|edges(P_1)| = 15$, $BExpr(P_1) = \{(lc0, x == 0), (lb0, y < x)\}$ and $exval(t, P_1) = \{(lb0, y < x), false\}$. Additionally P_1 is supplemented with variables $v, v_{in}, v_{x==0}$ and $v_{y<x}$ to program P'_1 , where the $range$ values can be computed accordingly. An inter-procedural conservative interval analysis [20, 21] can detect $range_l(v) = range_l(v_{in}) = \emptyset$ for all $l \in L' = \{lb0', lc2\}$ and $range_l(v) = \{0, 1\}$ for all $l \in labels(P'_1) \setminus L'$. This is used to compute δ_f^+ , δ_b^+ and δ_s^+ . The edges $(b2, b3), (b7, b9), (b9, b2), (b3, b10), (b3, b5) \in edges(P_1)$ are never executed, which is discovered by the interval analysis. This results in $\delta_d^+(P_1) = \frac{5}{15}$. Thus, the coverage metrics and gaps of Table 1 can be calculated for P_1 as described in the previous sections. The values were obtained using our current implementation of the program described in [8]. Computing the values presented in Table 1 takes less than two seconds on a modern Core i7 CPU equipped with 8 GiB RAM. Table 1 also contains the approximated values as presented in section 4. Although the coverage metrics are far less than 100 %, the test suite t is complete. Additional tests can not improve these coverages as confirmed by the gaps.

	$\gamma_f(t)$	δ_f^+	δ_f	δ_f^-	$\gamma_s(t)$	δ_s^+	δ_s	δ_s^-	$\gamma_d(t)$	δ_d^+	δ_d	δ_d^-
P_1	0.5	0.0	0.5	0.5	0.33	0.13	0.67	0.67	0.13	0.33	0.87	0.87
P_2	0.5	0.0	0.0	0.5	0.36	0.0	0.09	0.64	0.20	0.0	0.20	0.80
	$\gamma_b(t)$	δ_b^+	δ_b	δ_b^-	$\gamma_c(t)$	δ_c^+	δ_c	δ_c^-				
P_1	0.25	0.17	0.75	0.75	0.25	0.50	0.75	0.75				
P_2	0.38	0.0	0.13	0.62	0.75	0.0	0.13	0.25				

Table 1: Code Coverages and Gaps for P_1 in Fig. 1 and P_2 in Fig. 2

$label(P_2)$	$range_*^-(x)$	$range_*^-(y)$	$range_*^-(z)$	$range_*^-(w)$
m1,m2,m3,m5',lc0,lc1,lc2	\emptyset	\emptyset	\emptyset	\emptyset
m0,m4	$\{0, 1, 10\}$	$\{1, 5\}$	$\{0, 1, 10\}$	$\{1, 5\}$
m5,m6	$\{0\}$	$\{1, 5\}$	$\{0, 1, 10\}$	$\{1, 5\}$

Table 2: $range_l^-$ in P_2 using test suite t for P_2 in Fig. 2

The test suite t discovers $range_l^-(v) = \{0, 1\}$ for each reachable label l in t .

Most compilers, i.e. GCC and CL⁶ from Microsoft, are not able to do a flow-sensitive, context-sensitive inter-procedural analysis needed for a more precise lower bound in this example. The abstract interpretations done in a compiler or analysis tool do not yield such a precise lower bound, as most other tools are essentially model-checkers. Hence, the lower bound on the range for x and y would include all possible values at label $lb1$ in P_1 . Thus the lower bound on the function gap would be 0.

Additionally, using the tool gcov to compute the coverage of the test suite of example P_1 , no coverage is achieved by any test suite, because gcov does not take arithmetic exceptions into account resulting in

⁶Shipped with Microsoft Visual Studio

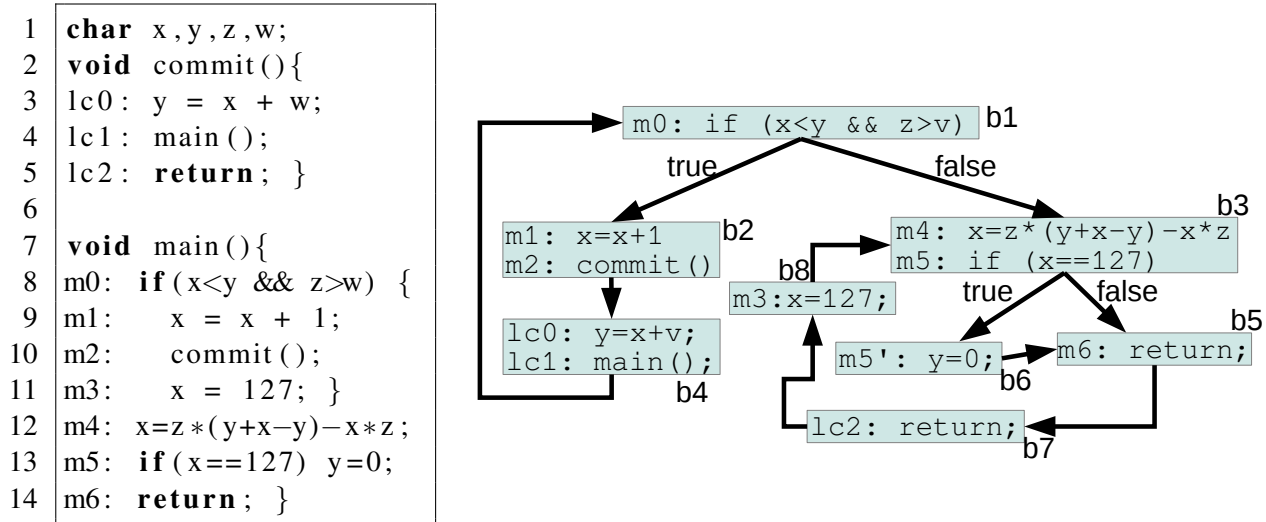


Figure 2: SPDS example P_2 in ISO-C syntax for dead code by bad condition + corresponding BBICFG

0% coverage. Of more practical relevance is the calculation of the coverage gap for non-arithmetic errors. For instance example P_2 in Fig. 2 has a difficult condition $(x < y \ \&\& \ z > w)$. The variables x , y , z and v of P_2 (Fig. 2) are global variables of type $char = bits(..) = 8$. The whole block below ($m1 - m3$) becomes dead, if the condition on label $m0$ evaluates to $false$. Thus `commit` is not called and the (indirect) recursion not started. Additionally, for all possible test cases, the condition $(x == 127)$ on label $m5$ never evaluates to $true$.

Let $t = \{(0, 1, 0, 1), (1, 1, 1, 1), (10, 5, 10, 5)\}$ be a test suite with (x, y, z, w) being the values set before calling `main`. It is $|stats(P_2)| = 11$, $|blocks(P_2)| = 8$, $|edges(P_2)| = 10$, $BExpr(P_2) = \{(m0, x < y), (m0, z > w), (m0, x < y \ \&\& \ z > w), (m5, x == 127)\}$ and $exval(t, P_2) = \{((m0, x < y), true), ((m0, x < y), false), ((m0, z > w), true), ((m0, z > w), false), ((m0, x < y \ \&\& \ z > w), false), ((m5, x == 127), false)\}$. The test $\alpha = (0, 1, 1, 0)$ would be a good candidate for the test suite t , because $\gamma_s(\{\alpha\}) = 91\%$ is perfect (proofed by the gap δ_s). Table 1 contains the calculated coverage metrics and gaps for P_2 . As one can see from the third line of Table 1, the exact gap in the existing code is rather small: it consists of the condition $x == 127$ on label $m5$ and the following code block. This is one of the examples in which our method can instruct the tester to expand the test suite. More code can not be covered, because in each test the variables x and z as well as y and v are aliases. Although most of the code in the example is alive. As seen in the previous example, the approximated lower and upper bounds are not perfect. An upper bound on the gap δ_f of called functions, is $\delta_f^- = 1 - 0.5$, whereas from the two available functions one was called during the execution of test suite t . It is $range_l(v) = \{0, 1\}$ for every label $l \in labels(P_2')$ in P_2' (supplementation of P_2). P_2 is also supplemented with variables $v_{in}, v_{x==0}$ and $v_{y < x}$, so that the $range$ values can be computed and approximated using an inter-procedural conservative interval analysis ($range^+$). The results for the variables are shown in Table 2 and 1.

Contrary to `gcov` the computation of the code coverages followed the C Program and did not rely on any symbolic assembler. Such abstractions might cause more coverage shown than the actual coverage in ISO-C. The statement coverage reported by `gcov` corresponds to γ_s . Values close or exactly corresponding to γ_X can be obtained from `gcov` for these particular examples. Not all values will match γ_X , because `gcov` uses a different definition for decision and branch coverage and relies on symbolic assembler output.

6 Related Work

A lot of research is done to get a better coverage for a test suite [4]. However an important point is often missing: often it is impossible to cover 100% of the code in practice, because of gaps.

To the best of our knowledge no research has been done to compute provable exact gaps used in code coverage. Conservative strategies underestimate the coverage gap [5]. Current research only approximates the gap. For instance [3] presents a method to automatically add tests by computing a gap of code covered by the test suite and possible code coverage. The authors miss the important point of having code which will and can never be used. In [3] the emphasis is on large scale projects, but especially in such large projects there is code which cannot be executed and should be removed by the compiler. As [11] describes, some code parts are more important than others. Testing parts of a program which will never be executed is then a loss of resources. Gittens et al. use a domain expert to categorize the code, i.e. for which parts of the source code their tool should generate tests automatically. Our gap computation presented in this paper could be used to automatically categorize the code and not depend on a domain expert. Another project of interest is [16] by Kicillof et al., which shows how to create checkable models. The focus of Kicillof et al. are models which can be created by stakeholders or maybe even marketing experts, and thus is directed at their specific problems at Microsoft. Most other research concerning the computation of gaps in coverage targets the pre-silicon design validation, i.e. [6, 19].

Both papers on pre-silicon design validation are not concerned with testing gaps. They rather check if a specification can be achieved. However, our paper is concerned with languages similar to C, not any Register Transfer Language (RTL) or even specifications.

As Regehr correctly writes in [15], such specifications, which are checked in [6, 19] might have been wrong in the first place. One solution proposed by Regehr for finding errors in specifications is having more people to look over these. A different solution uses our tool and computes parts of the realized specification that are never used, thus giving hints to erroneous specifications.

Whereas Berner et al. are targeting the user of an automatic test system [4] our method targets the automatic test system itself. Berner et al. describe *lessons learned* from their experience with code coverage analysis tools and automatic test generation tools and propose a list of rules to be followed when introducing and using an automatic test tool. Our research was not concerned with usability and group dynamics in a programming environment.

To the best of our knowledge the current research in testing, be it concolic⁷ or model-based, is not concerned with the actual problems of code coverage gaps. Gap coverage analysis is not only useful in test case generation but also in the verification of functional correctness. Imagine the case of a dead function granting more user rights, it is easy to use a buffer overflow to trigger this functionality. Similar methods have been used by the CCC for analyzing and using a trojan horse [10]⁸.

Another important tool, which might be able to compete with our method is Frama-C [5]⁹. Frama-C is a conservative analysis tool, which is able to find dead code, execute a static value analysis and, contrary to gcov, is able to detect runtime-errors triggered for instance by division-by-zero. One of the differences between Frama-C and the method we propose in this paper is the theory behind it. In contrast to Frama-C [5] our method uses exact computation, does not overapproximate the values and does not rely on an experienced user. Our exact value analysis produces neither false negatives nor false positives as in Frama-C. Although their value analysis sometimes detects that a function does not terminate, it cannot be used to prove that a function terminates in general.

⁷interwoven concrete and symbolic execution

⁸especially the section *Upload- und Execute-Mechanismus*

⁹<http://frama-c.com>

Frama-C provides sophisticated plugins, but not all of them handle recursion properly. No sophisticated examples can be handled by Frama-C's value analysis. Some of the examples tested even cause runtime-errors in Frama-C itself, thus it is not reliable¹⁰.

As our review of the research indicates, none of current research done in testing is concerned with *exact* gap computation.

7 Summary and Conclusions

This paper presents a framework to compute exact gaps between the feasible and theoretical maximal possible code coverage value. For specifying programs in an ISO-C semantic we use a very powerful model, namely SPDS. The power of SPDS allows to model an ISO-C compatible semantics for programs without abstraction. Therefore we are able to do an exact value analysis using model checking techniques and so we obtain exact gaps. We describe how to efficiently approximate the gap in all the other cases. When using flow-sensitive, path-sensitive, inter-procedural and context-sensitive data flow analyses for approximating the exact values one can also use a model-checking tool. The biggest problems of using a model-checker are false positives or false negatives caused by abstraction. Instead, our approach does not rely on such heavy abstraction and does not cause false alarms on our ISO-C compatible semantic. Thus user input or feedback is not required to decide about false alarms. A lot of computing power is required for using such powerful models. Due to smaller programs and smaller data types our approach is still practical for embedded systems.

Having combined the best parts of model-checking and static analyses we use expansive model-checking only when needed (e.g. the gap approximation bounds are not small enough). Thus the computation of $Post^*$ is needed only if the gap approximation using static analysis and a test suite is not exact ($\delta^- \neq \delta^+$).

Using our method a lot of metrics can be compared better among each other now, because of exactly specified gaps. Our method allows the testing of non-functional requirements, too. For example the worst case execution times (WCET) using a WCET metric¹¹ can be computed.

Our current research considers the practical relevance of exact gap computation for verification of software especially in the area of compiler correctness. Additionally we are considering other areas of research to apply the computation of exact values and exact gaps. For example the computation of exact value ranges can be used for verification of components [2].

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¹⁰It should be noted, that these runtime-errors should vanish in future versions

¹¹e.g. $\gamma_{WCET}(t) := \max_{a \in t} runtime(a)$ with supplemetation $tick = tick + 1$ on each statement, such that $\gamma_{WCET}(t) = \max(\cup_{l \in labels(S)} range_l(tick))$.

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