# A bytecode logic for JML and types (Isabelle/HOL sources) 

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#### Abstract

This document contains the Isabelle/HOL sources underlying our paper A bytecode logic for JML and types [2], updated to Isabelle 2008. We present a program logic for a subset of sequential Java bytecode that is suitable for representing both, features found in high-level specification language JML as well as interpretations of high-level type systems. To this end, we introduce a fine-grained collection of assertions, including strong invariants, local annotations and VDM-reminiscent partial-correctness specifications. Thanks to a goal-oriented structure and interpretation of judgements, verification may proceed without recourse to an additional control flow analysis. The suitability for interpreting intensional type systems is illustrated by the proof-carryingcode style encoding of a type system for a first-order functional language which guarantees a constant upper bound on the number of objects allocated throughout an execution, be the execution terminating or non-terminating.

Like the published paper, the formal development is restricted to a comparatively small subset of the JVML, lacking (among other features) exceptions, arrays, virtual methods, and static fields. This shortcoming has been overcome meanwhile, as our paper has formed the basis of the Mobius base logic [9], a program logic for the full sequential fragment of the JVML. Indeed, the present formalisation formed the basis of a subsequent formalisation of the Mobius base logic in the proof assistant Coq, which includes a proof of soundness with respect to the Bicolano operational semantics [10].


## Contents

1 Preliminaries: association lists 2
2 Language 3
2.1 Syntax . . . . . . . . . . . . . . . . . . . . . . . . . . . . . . . 3
2.2 Dynamic semantics . . . . . . . . . . . . . . . . . . . . . . . . 5
2.2.1 Semantic components . . . . . . . . . . . . . . . . . . 5
2.2.2 Operational judgements ..... 6
2.3 Basic properties ..... 8
3 Axiomatic semantics ..... 8
3.1 Assertion forms ..... 8
3.2 Proof system ..... 9
4 Auxiliary operational judgements ..... 14
4.1 Multistep execution ..... 14
4.2 Reachability relation ..... 15
5 Soundness ..... 16
5.1 Validity ..... 16
5.2 Soundness under valid contexts ..... 18
5.3 Soundness of verified programs ..... 20
6 A derived logic for a strong type system ..... 20
6.1 Syntax and semantics of judgements ..... 21
6.2 Derived proof rules ..... 22
6.3 Soundness of high-level type system ..... 23

## 1 Preliminaries: association lists

Finite maps are used frequently, both in the representation of syntax and in the program logics. Instead of restricting Isabelle's partial map type $\alpha \rightharpoonup \beta$ to finite domains, we found it easier for the present development to use the following adhoc data type of association lists.
type-synonym ( $\left.{ }^{\prime} a,{ }^{\prime} b\right)$ AssList $=\left({ }^{\prime} a \times\right.$ ' $\left.b\right)$ list
primrec lookup::('a, 'b) AssList $\Rightarrow{ }^{\prime} a \Rightarrow(' b$ option $)(-\downarrow-[90,0] 90)$
where
lookup [] $l=$ None |
lookup $(h \# t) l=($ if fst $h=l$ then Some $($ snd $h)$ else lookup $t l)$
$\langle p r o o f\rangle\langle p r o o f\rangle\langle p r o o f\rangle\langle p r o o f\rangle\langle$ proof $\rangle$
The statement following the type declaration of lookup indicates that we may use the infix notation $L \downarrow a$ for the lookup operation, and asserts some precedence for bracketing. In a similar way, shorthands are introduced for various operations throughout this document.
primrec delete::('a, 'b) AssList $\Rightarrow{ }^{\prime} a \Rightarrow\left({ }^{\prime} a, ~ ' b\right)$ AssList
where
delete [] $a=[] \mid$
delete $(h \# t) a=($ if $(f s t h)=a$ then delete $t$ a else $(h \#($ delete $t a)))$
$\langle$ proof $\rangle\langle$ proof $\rangle\langle$ proof $\rangle\langle$ proof $\rangle\langle$ proof $\rangle\langle$ proof $\rangle\langle$ proof $\rangle$
definition upd::('a, 'b) AssList $\Rightarrow^{\prime} a \Rightarrow{ }^{\prime} b \Rightarrow\left({ }^{\prime} a,{ }^{\prime} b\right)$ AssList
$(-[-\rightarrow-][1000,0,0] 1000)$

The empty map is represented by the empty list.

```
definition emp::('a, 'b)AssList
where emp \(=[]\langle\) proof \(\rangle\)
definition contained \(::(\) ('a,'b) AssList \(\Rightarrow\) ('a,'b) AssList \(\Rightarrow\) bool
where contained \(L M=(\forall a b . L \downarrow a=\) Some \(b \longrightarrow M \downarrow a=\) Some \(b)\)
```

The following operation defined the cardinality of a map.

```
fun AL-Size :: ('a, 'b) AssList \(\Rightarrow\) nat (|-| [1000] 1000) where
    AL-Size [] \(=0\)
\(\mid \operatorname{AL-Size}(h \# t)=\operatorname{Suc}(A L\)-Size \((\operatorname{delete} t(f s t h)))\)
\(\langle p r o o f\rangle\langle p r o o f\rangle\langle p r o o f\rangle\langle p r o o f\rangle\langle p r o o f\rangle\langle p r o o f\rangle\)
```

Some obvious basic properties of association lists and their operations are easily proven, but have been suppressed during the document preparation.

## 2 Language

### 2.1 Syntax

We have syntactic classes of (local) variables, class names, field names, and method names. Naming restrictions, namespaces, long Java names etc. are not modelled.
typedecl Var
typedecl Class
typedecl Field
typedecl Method
Since arithmetic operations are modelled as unimplemented functions, we introduce the type of values in this section. The domain of heap locations is arbitrary.

## typedecl $A d d r$

A reference is either null or an address.
datatype Ref $=$ Nullref $\mid$ Loc Addr
Values are either integer numbers or references.
datatype Val $=$ RVal Ref $\mid$ IVal int
The type of (instruction) labels is fixed, since the operational semantics increments the program counter after each instruction.
type-synonym Label $=$ int

Regarding the instructions, we support basic operand-stack manipulations, object creation, field modifications, casts, static method invocations, conditional and unconditional jumps, and a return instruction.
For every (Isabelle) function $f: V a l \Rightarrow V a l \Rightarrow V a l$ we have an instruction binop $f$ whose semantics is to invoke $f$ on the two topmost values on the operand stack and replace them with the result. Similarly for unop $f$.

```
datatype Instr \(=\)
    const Val
| dup
| pop
| swap
| load Var
| store Var
binop Val \(\Rightarrow\) Val \(\Rightarrow\) Val
| unop Val \(\Rightarrow\) Val
| new Class
| getfield Class Field
| putfield Class Field
| checkcast Class
| invokeS Class Method
| goto Label
| iftrue Label
| vreturn
```

Method body declarations contain a list of formal parameters, a mapping from instruction labels to instructions, and a start label. The operational semantics assumes that instructions are labelled consecutively ${ }^{1}$.
type-synonym Mbody $=$ Var list $\times($ Label, Instr $)$ AssList $\times$ Label
A class definition associates method bodies to method names.
type-synonym Classdef $=($ Method, Mbody $)$ AssList
Finally, a program consists of classes.
type-synonym Prog $=($ Class, Classdef $)$ AssList
Taken together, the three types Prog, Classdef, and Mbody represent an abstract model of the virtual machine environment. In our opinion, it would be desirable to avoid modelling this environment at a finer level, at least for the purpose of the program logic. For example, we prefer not to consider in detail the representation of the constant pool.

[^0]
### 2.2 Dynamic semantics

### 2.2.1 Semantic components

An object consists of the identifier of its dynamic class and a map from field names to values. Currently, we do not model type-correctness, nor do we require that all (or indeed any) of the fields stem from the static definition of the class, or a super-class. Note, however, that type correctness can be expressed in the logic.
type-synonym Object $=$ Class $\times($ Field, Val $)$ AssList
The heap is represented as a map from addresses to values. The JVM specification does not prescribe any particular object layout. The proposed type reflects this indeterminacy, but allows one to calculate the byte-correct size of a heap only after a layout scheme has been supplied. Alternative heap models would be the store-less semantics in the sense of Jonkers [8] and Deutsch [5], (where the heap is modelled as a partial equivalence relation on access paths), or object-based semantics in the sense of Reddy [11], where the heap is represented as a history of update operations. Hähnle et al. use a variant of the latter in their dynamic logic for a JavaCard [6].
type-synonym Heap $=($ Addr, Object $)$ AssList
Later, one might extend heaps by a component for static fields.
The types of the (register) store and the operand stack are as expected.
type-synonym Store $=($ Var, Val $)$ AssList
type-synonym OpStack $=$ Val list
States contain an operand stack, a store, and a heap.

```
type-synonym State = OpStack }\times\mathrm{ Store }\times\mathrm{ Heap
definition heap::State }=>\mathrm{ Heap
where heap s=\operatorname{snd}(\mathrm{ snd s)}
```

The operational semantics and the program logic are defined relative to a fixed program $P$. Alternatively, the type of the operational semantics (and proof judgements) could be extended by a program component. We also define the constant value $T R U E$, the representation of which does not matter for the current formalisation.

```
axiomatization P::Prog and TRUE::Val
```

In order to obtain more readable rules, we define operations for extracting method bodies and instructions from the program.

```
definition mbody-is::Class \(\Rightarrow\) Method \(\Rightarrow\) Mbody \(\Rightarrow\) bool
where mbody-is \(C m M=(\exists C D . P \downarrow C=\) Some \(C D \wedge C D \downarrow m=\) Some \(M)\)
```

```
definition get-ins::Mbody \(\Rightarrow\) Label \(\Rightarrow\) Instr option
where get-ins \(M l=(f s t(\operatorname{snd} M)) \downarrow l\)
definition ins-is::Class \(\Rightarrow\) Method \(\Rightarrow\) Label \(\Rightarrow\) Instr \(\Rightarrow\) bool
where ins-is \(C m\) l ins \(=(\exists M . \operatorname{mbody-is} C m M \wedge\) get-ins \(M l=\) Some ins \()\)
```

The transfer of method arguments from the caller's operand stack to the formal parameters of an invoked method is modelled by the predicate

```
inductive-set Frame::(OpStack }\times(\mathrm{ Var list ) }\times\mathrm{ Store }\times\mathrm{ OpStack) set
where
FrameNil: \llbracketoo=ops\rrbracket \Longrightarrow(ops,\],emp,oo): Frame
Frame-cons: \llbracket(oo,par,S,ops) : Frame; R=S[x\mapstov]\rrbracket
    "(v # oo, x # par,R,ops):Frame
<proof\rangle
```

In order to obtain a deterministic semantics, we assume the existence of a function, with the obvious freshness axiom for this construction.
axiomatization nextLoc::Heap $\Rightarrow A d d r$
where nextLoc-fresh: $h \downarrow$ (nextLoc $h)=$ None

### 2.2.2 Operational judgements

Similar to Bannwart-Müller [1], we define two operational judgements: a one-step relation and a relation that represents the transitive closure of the former until the end of the current method invocation. These relations are mutually recursive, since the method invocation rule contracts the execution of the invoked method to a single step. The one-step relation associates a state to its immediate successor state, where the program counter is interpreted with respect to the current method body. The transitive closure ignores the bottom part of the operand stack and the store of the final configuration. It simply returns the heap and the result of the method invocation, where the latter is given by the topmost value on the operand stack. In contrast to [1], we do not use an explicit return variable. Both relations take an additional index of type nat that monitors the derivation height. This is useful in the proof of soundness of the program logic.

Intuitively, ( $M, l, s, n, l^{\prime}, s^{\prime}$ ):Step means that method (body) $M$ evolves in one step from state $s$ to state $s^{\prime}$, while statement $(M, s, n, h, v)$ :Exec indicates that executing from $s$ in method $M$ leads eventually to a state whose final value is $h$, where precisely the last step in this sequence is a vreturn instruction and the return value is $v$.

Like Bannwart and Müller, we define a "frame-less" semantics. i.e. the execution of a method body is modelled by a transitive closure of the basic step-relation, which results in a one-step reduction at the invocation site.

Arguably, an operational semantics with an explicit frame stack is closer to the real JVM. It should not be difficult to verify the operational soundness of the present system w.r.t. such a finer model, or to modify the semantics.

```
inductive-set
    Step::(Mbody }\times\mathrm{ Label }\times\mathrm{ State }\times\mathrm{ nat }\times\mathrm{ Label }\times\mathrm{ State) set
and
    Exec::(Mbody }\times\mathrm{ Label }\times\mathrm{ State }\times\mathrm{ nat }\times\mathrm{ Heap }\times\mathrm{ Val) set
where
Const:\llbracketget-ins M l = Some (const v);NEXT = (v # os,s,h); ll=l+1\rrbracket
        \Longrightarrow ( M , l , ( o s , s , h ) , 1 , l l , N E X T ) : S t e p
|
Dup: \llbracketget-ins M l = Some dup; NEXT = (v#v# os,s,h); ll=l+1\rrbracket
    \Longrightarrow ( M , l , ( v \# o s , s , h ) , 1 , l l , N E X T ) : S t e p
|
Pop: \llbracketget-ins M l = Some pop; NEXT = (os,s,h);ll=l+1\rrbracket
    \Longrightarrow ( M , l , ( v \# o s , s , h ) , 1 , l l , N E X T ) : S t e p
|
Swap:\llbracketget-ins M l = Some swap;NEXT = (w# (v# os),s,h);ll=l+1\rrbracket
    \Longrightarrow(M,l,(v# (w# os),s,h), 1,ll,NEXT):Step
|
Load:\llbracketget-ins Ml=Some (load x); s\downarrowx= Some v;
    NEXT = (v# os,s,h);ll=l+1\rrbracket
    \Longrightarrow ( M , l , ( o s , s , h ) , 1 , l l , N E X T ) : S t e p
|
Store:\llbracketget-ins M l = Some (store x);NEXT = (os,s[x\mapstov],h);ll=l+1\rrbracket
    \Longrightarrow ( M , l , ( v \# o s , s , h ) , 1 , l l , N E X T ) : S t e p
|
Binop:\llbracketget-ins M l = Some (binop f); NEXT = ((fv w) # os,s,h); ll=l+1\rrbracket
    \Longrightarrow(M,l,(v# (w # os),s,h),1,ll,NEXT) : Step
Unop:\llbracketget-ins M l = Some (unop f);NEXT = ((f v)# os,s,h);ll=l+1\rrbracket
    \Longrightarrow ( M , l , ( v \# o s , s , h ) , 1 , l l , N E X T ) : S t e p
|
New: \llbracketget-ins M l = Some (new d); newobj = (d, emp); a=nextLoc h;
        NEXT = ((RVal (Loc a)) # os,s,h[a\mapstonewobj]);ll=l+1]
    \Longrightarrow ( M , l , ( o s , s , h ) , 1 , l l , N E X T ) : S t e p
|
Get: \llbracketget-ins M l = Some (getfield d F); h\downarrowa=Some (d, Flds);
        Flds\downarrowF=Some v; NEXT = (v # os,s,h); ll=l+1\rrbracket
    \Longrightarrow(M,l,((RVal (Loc a)) # os,s,h), 1, ll,NEXT) : Step
Put: \llbracketget-ins M l = Some (putfield d F); h\downarrowa=Some (d, Flds);
    newobj = (d, Flds[F\mapstov]);NEXT = (os,s,h[a\mapstonewobj]);ll=l+1\rrbracket
    \Longrightarrow(M,l,(v # ((RVal (Loc a)) # os),s,h), 1, ll,NEXT) : Step
|
Cast: \llbracketget-ins M l = Some (checkcast d); h\downarrowa=Some (d, Flds);
    NEXT = ((RVal (Loc a)) # os,s,h);ll=l+1\rrbracket
    \Longrightarrow ( M , l , ( ( R V a l ~ ( L o c ~ a ) ) ~ \# ~ o s , s , h ) , ~ 1 , ~ l l , N E X T ) ~ : ~ S t e p ~
|
```

```
Goto: \(\llbracket\) get-ins \(M l=\) Some \((\) goto \(p c) \rrbracket \Longrightarrow(M, l, S, 1, p c, S):\) Step
IfT: \(\llbracket\) get-ins \(M l=\) Some \((\) iftrue \(p c) ; N E X T=(o s, s, h) \rrbracket\)
    \(\Longrightarrow(M, l,(T R U E \#\) os \(, s, h), 1, p c, N E X T): S t e p\)
\(\mid\)
IfF: \(\llbracket\) get-ins \(M l=\) Some (iftrue \(p c) ; v \neq T R U E ; N E X T=(o s, s, h) ; l l=l+1 \rrbracket\)
    \(\Longrightarrow(M, l,(v \# o s, s, h), 1, l l, N E X T):\) Step
|
InvS: \(\llbracket\) get-ins Mlosome (invokeS Cm); mbody-is C m (par,code,l0);
    ((par,code,l0),l0,([], S,h), n, hh, v): Exec;
    (ops,par,S,os) : Frame; NEXT = \((v \#\) os,s,hh); ll=l+1】
    \(\Longrightarrow(M, l,(o p s, s, h), S u c n, l l, N E X T): S t e p\)
|
Vret: \(\llbracket\) get-ins \(M l=\) Some vreturn \(\rrbracket(M, l,(v \#\) os \(, s, h), 1, h, v):\) Exec
|
Run: \(\llbracket(M, l, s, n, l l, t):\) Step \(;(M, l l, t, m, h, v): E x e c ; k=(\max n m)+1 \rrbracket\)
    \(\Longrightarrow(M, l, s, k, h, v):\) Exec
```

A big-step operational judgement that abstracts from the derivation height is easily defined.
definition Opsem $::$ Mbody $\Rightarrow$ Label $\Rightarrow$ State $\Rightarrow$ Heap $\Rightarrow$ Val $\Rightarrow$ bool where Opsem Mlshv=( $\exists$ n. $(M, l, s, n, h, v)$ :Exec $)$

### 2.3 Basic properties

We provide elimination lemmas for the inductively defined relations

## inductive-cases eval-cases:

```
(M,l,s,n,ll,t) : Step
(M,l,s,n,h,v) : Exec\langleproof\rangle\langleproof\rangle
```

and observe that no derivations of height 0 exist.
lemma no-zero-height-Step-derivs: $(M, l, s, 0, l l, t):$ Step $\Longrightarrow$ False $\langle$ proof $\rangle\langle p r o o f\rangle$
lemma no-zero-height-Exec-derivs: $(M, l, s, 0, h, v):$ Exec $\Longrightarrow$ False $\langle$ proof $\rangle\langle$ proof $\rangle\langle$ proof $\rangle\langle$ proof $\rangle\langle$ proof $\rangle\langle$ proof $\rangle\langle p$
By induction on the derivation system one can show determinism.
$\langle p r o o f\rangle$
lemma Step-determ:
$\llbracket(M, l, s, n, l 1, t) \in \operatorname{Step} ;(M, l, s, m, l 2, r): S t e p \rrbracket \Longrightarrow n=m \wedge t=r \wedge l 1=l 2\langle p r o o f\rangle$
lemma Exec-determ:
$\llbracket(M, l, s, n, h, v) \in E x e c ;(M, l, s, m, k, w): E x e c \rrbracket \Longrightarrow n=m \wedge h=k \wedge v=w\langle p r o o f\rangle$

## 3 Axiomatic semantics

### 3.1 Assertion forms

We introduce two further kinds of states. Initial states do not contain operand stacks, terminal states lack operand stacks and local variables, but include return values.

```
type-synonym InitState = Store }\times\mathrm{ Heap
type-synonym TermState = Heap }\times\mathrm{ Val
```

A judgements relating to a specific program point C.m.l consists of pre- and post-conditions, an invariant, and - optionally - a local annotation. Local pre-conditions and annotations relate initial states to states, i.e. are of type
type-synonym Assn $=$ InitState $\Rightarrow$ State $\Rightarrow$ bool
Post-conditions additionally depend on a terminal state
type-synonym Post $=$ InitState $\Rightarrow$ State $\Rightarrow$ TermState $\Rightarrow$ bool
Invariants hold for the heap components of all future (reachable) states in the current frame as well as its subframes. They relate these heaps to the current state and the initial state of the current frame.
type-synonym Inv $=$ InitState $\Rightarrow$ State $\Rightarrow$ Heap $\Rightarrow$ bool
Local annotations of a method implementation are collected in a table of type
type-synonym ANNO $=($ Label, Assn $)$ AssList
Implicitly, the labels are always interpreted with respect to the current method. In addition to such a table, the behaviour of methods is specified by a partial-correctness assertion of type
type-synonym MethSpec $=$ InitState $\Rightarrow$ TermState $\Rightarrow$ bool
and a method invariant of type
type-synonym MethInv $=$ InitState $\Rightarrow$ Heap $\Rightarrow$ bool
A method invariant is expected to be satisfied by the heap components of all states during the execution of the method, including states in subframes, irrespectively of the termination behaviour.

All method specifications are collected in a table of type
type-synonym MSPEC $=($ Class $\times$ Method, MethSpec $\times$ MethInv $\times$ ANNO $)$ AssList

A table of this type assigns to each method a partial-correctness specification, a method invariant, and a table of local annotations for the instructions in this method.

### 3.2 Proof system

The proof system derives judgements of the form $G \triangleright\{A\} C, m, l\{B\} I$ where $A$ is a pre-condition, $B$ is a post-condition, $I$ is a strong invariant, C.m.l represents a program point, and $G$ is a proof context (see below). The proof
system consists of syntax-directed rules and structural rules. These rules are formulated in such a way that assertions in the conclusions are unconstrained, i.e. a rule can be directly applied to derive a judgement. In the case of the syntax-directed rules, the hypotheses require the derivation of related statements for the control flow successor instructions. Judgements occurring as hypotheses involve assertions that are notationally constrained, and relate to the conclusions' assertions via uniform constructions that resemble strongest postconditions in Hoare-style logics.

On pre-conditions, the operator
definition SP-pre::Mbody $\Rightarrow$ Label $\Rightarrow$ Assn $\Rightarrow$ Assn
where SP-pre M l $A=(\lambda s 0 r .(\exists s l 1 n . A s 0 s \wedge(M, l, s, n, l 1, r): S t e p))$
constructs an assertion that holds of a state $r$ precisely if the argument assertion $A$ held at the predecessor state of $r$. Similar readings explain the constructions on post-conditions
definition SP-post::Mbody $\Rightarrow$ Label $\Rightarrow$ Post $\Rightarrow$ Post
where $S P$-post $M l B=(\lambda s 0 r t .(\forall s l 1 n .(M, l, s, n, l 1, r): S t e p \longrightarrow B s 0 s t))$
and invariants
definition $S P$-inv:: Mbody $\Rightarrow$ Label $\Rightarrow$ Inv $\Rightarrow$ Inv
where $S P-i n v M l I=(\lambda s 0 r h . \forall$ sl1 $n .(M, l, s, n, l 1, r): S t e p \longrightarrow I s 0$ sh)
For the basic instructions, the appearance of the single-step execution relation in these constructions makes the strongest-postcondition interpretation apparent, but could easily be eliminated by unfolding the definition of Step. In the proof rule for static method invocations, such a direct reference to the operational semantics is clearly undesirable. Instead, the proof rule extracts the invoked method's specification from the specification table. In order to simplify the formulation of the proof rule, we introduce three operators which manipulate the extracted assertions in a similar way as the above $S P$-operators.
definition SINV-pre:: Var list $\Rightarrow$ MethSpec $\Rightarrow$ Assn $\Rightarrow$ Assn where SINV-pre par T $A=$
( $\lambda$ s0s. $(\exists$ ops1 ops2 $S R h k w$.
(ops1,par,R,ops2) : Frame $\wedge T(R, k)(h, w) \wedge$
A s0 $(o p s 1, S, k) \wedge s=(w \# o p s 2, S, h)))$
definition SINV-post:: Var list $\Rightarrow$ MethSpec $\Rightarrow$ Post $\Rightarrow$ Post where SINV-post par T $B=$
( $\lambda$ s0 st. $\forall$ ops1 ops2 $S R h k w r$.
(ops1, par, R,ops2) : Frame $\longrightarrow T(R, k)(h, w) \longrightarrow$
$s=(w \# o p s 2, S, h) \longrightarrow r=(o p s 1, S, k) \longrightarrow B s 0 r t)$
definition SINV-inv:: Var list $\Rightarrow$ MethSpec $\Rightarrow$ Inv $\Rightarrow$ Inv where
SINV-inv par $T I=$

$$
\begin{aligned}
&\left(\lambda s 0 s h^{\prime}\right. . \forall \text { ops1 ops2 } S R h k w . \\
&(o p s 1, \text { par }, R, o p s 2): \text { Frame } \longrightarrow T(R, k)(h, w) \longrightarrow \\
&\left.s=(w \# o p s 2, S, h) \longrightarrow I \text { s0 }(o p s 1, S, k) h^{\prime}\right)
\end{aligned}
$$

The derivation system is formulated using contexts $G$ of proof-theoretic assumptions representing local judgements. The type of contexts is
type-synonym CTXT $=($ Class $\times$ Method $\times$ Label, Assn $\times$ Post $\times$ Inv $)$ AssList
The existence of the proof context also motivates that the hypotheses in the syntax-directed rules are formulated using an auxiliary judgement form, $G \triangleright\langle A\rangle C, m, l\langle B\rangle I$. Statements for this auxiliary form can be derived by essentially two rules. The first rule, $A X$, allows us to extract assumptions from the context, while the second rule, $I N J E C T$, converts an ordinary judgement $G \triangleright\{A\} C, m, l\{B\} I$ into $G \triangleright\langle A\rangle C, m, l\langle B\rangle I$. No rule is provided for a direct embedding in the opposite direction. As a consequence of this formulation, contextual assumptions cannot be used directly to justify a statement $G \triangleright\{A\} C, m, l\{B\} I$. Instead, the extraction of such an assumption has to be followed by at least one "proper" (syntax-driven) rule. In particular, attempts to verify jumps by assuming a judgement and immediately using it to prove the rule's hypothesis are ruled out. More specifically, the purpose of this technical device will become obvious when we discharge the context in the proof of soundness (Section 5.3). Here, each entry $\left(\left(C^{\prime}, m^{\prime}, l^{\prime}\right),\left(A^{\prime}, B^{\prime}, I^{\prime}\right)\right)$ of the context will need to be justified by a derivation $G \triangleright\left\{A^{\prime}\right\} C^{\prime}, m^{\prime}, l^{\prime}\left\{B^{\prime}\right\} I^{\prime}$. This justification should in principle be allowed to rely on other entries of the context. However, an axiom rule for judgements $G \triangleright\{A\} C, m, l\{B\} I$ would allow a trivial discharge of any context entry. The introduction of the auxiliary judgement form $G \triangleright\langle A\rangle C, m, l\langle B\rangle I$ thus ensures that the discharge of a contextual assumption involves at least one application of a "proper" (i.e. syntax-directed) rule. Rule INJECT is used to chain together syntax-directed rules. Indeed, it allows us to discharge a hypothesis $G \triangleright\langle A\rangle C, m, l\langle B\rangle I$ of a syntax-directed rule using a derivation of $G \triangleright\{A\} C, m, l\{B\} I$.

The proof rules are defined relative to a fixed method specification table $M S T$.
axiomatization $M S T:: M S P E C$
In Isabelle, the distinction between the two judgement forms may for example be achieved by introducing a single judgement form that includes a boolean flag, and definiing two pretty-printing notions.

```
inductive-set \(S P\)-Judgement ::
    (bool \(\times\) CTXT \(\times\) Class \(\times\) Method \(\times\) Label \(\times\) Assn \(\times\) Post \(\times\) Inv) set
and
    SP-Deriv :: CTXT => Assn => Class \(=>\) Method \(=>\) Label \(=>\)
            Post \(=>\) Inv \(=>\) bool
```

```
    \((-\triangleright\{-\}-,-,-\{-\}-[100,100,100,100,100,100,100] 50)\)
and
    SP-Assum :: CTXT \(=>\) Assn \(=>\) Class \(=>\) Method \(=>\) Label \(=>\)
        Post \(=>\) Inv \(=>\) bool
    \((-\triangleright\langle-\rangle-,-,-\langle-\rangle-[100,100,100,100,100,100,100] 50)\)
where
\(G \triangleright\{A\} C, m, l\{B\} I==(\) False \(, G, C, m, l, A, B, I): S P\)-Judgement
|
\(G \triangleright\langle A\rangle C, m, l\langle B\rangle I==(\) True \(G, C, m, l, A, B, I): S P\)-Judgement
|
INSTR:
【 mbody-is C m M ; get-ins Ml=Some ins;
        lookup \(\operatorname{MST}(C, m)=\) Some (Mspec, Minv,Anno);
    \(\forall Q\). lookup Anno \(l=\) Some \(Q \longrightarrow(\forall\) s0 s.A s0 s \(\longrightarrow Q\) s0 s);
    \(\forall s 0 s . A s 0 s \longrightarrow I s 0 s(h e a p s) ;\)
    ins \(\in\{\) const \(c\), dup, pop, swap, load \(x\), store \(x\), binop \(f\), unop \(g\),
            new d, getfield d \(F\), putfield d \(F\), checkcast d\};
    \(G \triangleright\langle(S P\)-pre \(M l A)\rangle C, m,(l+1)\langle(S P\)-post MlB) \(\rangle(S P-i n v M l I) \rrbracket\)
\(\Longrightarrow G \triangleright\{A\} C, m, l\{B\} I\)
\(\mid\)
GOTO:
    【 mbody-is C m M; get-ins Ml=Some (goto pc);
        \(\operatorname{MST} \downarrow(C, m)=\) Some (Mspec,Minv,Anno);
    \(\forall Q . A n n o \downarrow(l)=\) Some \(Q \longrightarrow(\forall s 0 s . A s 0 s \longrightarrow Q s 0 s)\);
    \(\forall s 0 s . A s 0 s \longrightarrow I s 0 s(\) heap \(s)\);
    \(G \triangleright\langle S P\)-pre \(M l A\rangle C, m, p c\langle S P-p o s t M l B\rangle(S P-i n v M l I) \rrbracket\)
\(\Longrightarrow G \triangleright\{A\} C, m, l\{B\} I\)
|
IF:
    【 mbody-is \(C m\); get-ins Ml=Some (iftrue pc);
        \(\operatorname{MST} \downarrow(C, m)=\) Some (Mspec, Minv,Anno);
        \(\forall Q . A n n o \downarrow(l)=\) Some \(Q \longrightarrow(\forall s 0 s . A s 0 s \longrightarrow Q s 0 s) ;\)
        \(\forall s 0 s . A s 0 s \longrightarrow I s 0 s(\) heap \(s) ;\)
        \(G \triangleright\langle S P-p r e M l(\lambda s 0 s .(\forall\) ops \(S k . s=(T R U E \# o p s, S, k) \longrightarrow A s 0 s))\rangle\)
            \(C, m, p c\)
            〈 SP-post Ml( \(\lambda\) s0 st.
                            \((\forall\) ops \(S k . s=(\) TRUE\#ops \(, S, k) \longrightarrow B\) s0 st \()\) ) \(\rangle\)
            ( SP-inv Ml ( \(\lambda\) s0 st.
                \((\forall\) ops \(S k . s=(\) TRUE\#ops \(, S, k) \longrightarrow I\) s0 st)) );
        \(G \triangleright\langle S P\)-pre \(M l(\lambda s 0 s\).
                \((\forall\) ops \(S k v . s=(v \# o p s, S, k) \longrightarrow v \neq T R U E \longrightarrow A s 0 s))\rangle\)
            C, m, (l+1)
            〈 SP-post Ml \((\lambda\) s 0 st.
                \((\forall\) ops \(S k v . s=(v \# o p s, S, k) \longrightarrow v \neq T R U E \longrightarrow B\) s0 st \()\) ) \(\rangle\)
            ( SP-inv Ml ( \(\lambda\) so st.
                \((\forall\) ops \(S k v . s=(v \# o p s, S, k) \longrightarrow v \neq T R U E \longrightarrow I s 0 s t))) \rrbracket\)
\(\Longrightarrow G \triangleright\{A\} C, m, l\{B\} I\)
```

```
VRET:
    \llbracket mbody-is C m M; get-ins M l = Some vreturn;
    MST\downarrow (C,m) = Some (Mspec,Minv,Anno);
    \forall.Anno\downarrow (l) = Some }Q\longrightarrow(\forall s0 s.A s0 s \longrightarrowQ s0 s)
    v0s.A s0 s\longrightarrowIs0 s(heap s);
    *0s.A s0s\longrightarrow(\forallvops Sh.s=(v#ops,S,h)\longrightarrowB s0s (h,v))\rrbracket
\Longrightarrow G \triangleright \{ A \} C , m , l \| B \} I
|
INVS:
\llbracket mbody-is C m M; get-ins M l = Some (invokeS D m
        MST\downarrow (C,m) = Some (Mspec,Minv,Anno);
        MST\downarrow (D,m') = Some (T,MI,Anno2); mbody-is D m' (par,code,l0);
        \forall.Anno\downarrow (l) = Some Q \longrightarrow( }\forall\mathrm{ s0 s.A s0 s }\longrightarrowQ s0 s)
        \forall0 s.A s0 s\longrightarrowI s0 s(heap s);
        *0 ops1 ops2 S R kt.(ops1,par,R,ops2): Frame \longrightarrow
            A s0(ops1,S,k)\longrightarrowMI (R,k)t\longrightarrowI s0(ops1,S,k) t;
        G\triangleright\langleSINV-pre par T A }|,m,(l+1)\langleSINV-post par T B
            (SINV-inv par T I)]
    \Longrightarrow\triangleright\triangleright{A}C,m,l { B } I
|
CONSEQ:
\llbracket (b,G,C,m,l,AA,BB,II) \inSP-Judgement;
        \forall0 s.A s0 s \longrightarrowAA s0 s; \forall s0 st.BB s0 st\longrightarrowB s0 st;
        *O sk. II s0 sk\longrightarrowI sO sk\rrbracket
    \Longrightarrow ( b , G , C , m , l , A , B , I ) \in S P - J u d g e m e n t
INJECT: \llbracketG\triangleright{A}C,m,l{B} I\rrbracket\LongrightarrowG\triangleright\langleA\rangleC,m,l\langleB\rangleI
AX:
\llbracket G\downarrow(C,m,l)=Some (A,B,I);MST\downarrow (C,m)=Some (Mspec,Minv,Anno);
    \forall.Anno\downarrow }(l)=\mathrm{ Some }Q\longrightarrow(\forall s0 s.A s0s\longrightarrowQ s0 s)
    v0 s.A s0 s\longrightarrowIs0 s(heap s)\rrbracket
    \LongrightarrowG\triangleright\langleA\rangleC,m,l \B\rangleI
```

As a first consequence, we can prove by induction on the proof system that a derivable judgement entails its strong invariant and an annotation that may be attached to the instruction.

```
proof\ranglelemma AssertionsImplyMethInvariants:
    \llbracket G\triangleright{A}C,m,l \ B } I; A s0 s\rrbracket\LongrightarrowIs0s(heap s)\langleproof\rangle
lemma AssertionsImplyAnnoInvariants:
    \llbracketG\triangleright{A}C,m,l{B}I; MST\downarrow (C,m)=Some(Mspec,Minv,Anno);
        Anno\downarrow }(l)=\mathrm{ Some Q;A s0 s\ CQ s0 s\proof>
```

For verified programs, all preconditions can be justified by proof derivations, and initial labels of all methods (again provably) satisfy the method preconditions.
definition mkState::InitState $\Rightarrow$ State
where mkState s0 $=([$, fst $s 0$,snd s0 $)$

```
definition \(m k\) Post::MethSpec \(\Rightarrow\) Post
where mkPost \(T=(\lambda s 0 s t . s=m k S t a t e s 0 \longrightarrow T s 0 t)\)
definition mkInv::MethInv \(\Rightarrow\) Inv
where mkInv MI \(=(\lambda s 0 s t . s=m k S t a t e s 0 \longrightarrow M I s 0 t)\)
definition \(V P-G:: C T X T \Rightarrow\) bool where
\(V P-G G=\)
    \(((\forall C m l A B I . G \downarrow(C, m, l)=\operatorname{Some}(A, B, I) \longrightarrow G \triangleright\{A\} C, m, l\{B\} I) \wedge\)
    ( \(\forall\) C m par code l0 T MI Anno.
    mbody-is C m (par,code,l0) \(\longrightarrow M S T \downarrow(C, m)=\operatorname{Some}(T, M I, A n n o) \longrightarrow\)
    \(G \triangleright\{(\lambda s 0 \mathrm{~s} . s=m k S t a t e ~ s 0)\} C, m, l 0\{m k P o s t T\}(m k I n v M I)))\)
definition \(V P::\) bool where \(V P=(\exists G . V P-G G)\)
```


## 4 Auxiliary operational judgements

Beside the basic operational judgements Step and Exec, the interpretation of judgements refers to two multi-step relations which we now define.

### 4.1 Multistep execution

The first additional operational judgement is the reflexive and transitive closure of Step. It relates states $s$ and $t$ if the latter can be reached from the former by a chain if single steps, all in the same frame. Note that $t$ does not need to be a terminal state. As was the case in the definition of Step, we first define a relation with an explicit derivation height index (MStep).

```
inductive-set
MStep \(::(\) Mbody \(\times\) Label \(\times\) State \(\times\) nat \(\times\) Label \(\times\) State \()\) set
where
\(M S\)-zero: \(\llbracket k=0 ; t=s ; l l=l \rrbracket \Longrightarrow(M, l, s, k, l l, t): M S t e p\)
|
MS-step: \(\llbracket(M, l, s, n, l 1, r): S t e p ;(M, l 1, r, k, l 2, t): M S t e p ; ~ m=S u c ~ k+n \rrbracket\)
    \(\Longrightarrow(M, l, s, m, l 2, t):\) MStep
```

The following properties of MStep are useful to notice.
$\langle p r o o f\rangle$ lemma ZeroHeightMultiElim: $(M, l, s, 0, l l, r) \in M S t e p \Longrightarrow r=s \wedge l l=l\langle p r o o f\rangle\langle p r o o f\rangle$
lemma MultiSplit:
$\llbracket(M, l, s, k, l l, t) \in$ MStep $; 1 \leq k \rrbracket \Longrightarrow$
$\exists n m r l 1 .(M, l, s, n, l 1, r): S t e p \wedge(M, l 1, r, m, l l, t): M S t e p \wedge S u c m+n=k\langle p r o o f\rangle\langle p r o o f\rangle$
lemma MStep-returnElim:
$\llbracket(M, l, s, k, l l, t) \in$ MStep $;$ get-ins $M l=$ Some vreturn $\rrbracket \Longrightarrow t=s \wedge l l=l\langle p r o o f\rangle\langle p r o o f\rangle$
lemma MultiApp:
$\llbracket(M, l, s, k, l 1, r): M S t e p ;(M, l 1, r, n, l 2, t): S t e p \rrbracket \Longrightarrow(M, l, s, S u c k+n, l 2, t): M S t e p\langle p r o o f\rangle\langle p r o o f\rangle$

## lemma MStep-Compose:

$\llbracket(M, l, s, n, l 1, r): M S t e p ;(M, l 1, r, k, l 2, t): M S t e p ; n k=n+k \rrbracket$
$\Longrightarrow(M, l, s, n k, l 2, t): M S t e p\langle$ proof $\rangle$
Here are two simple lemmas relating the operational judgements.

```
<proof\lemma MStep-Exec1:
    \llbracket(M,l,s,kb,l1,t)\inMStep; (M,l,s,k,hh,v)\inExec\rrbracket
    \Longrightarrow \exists ~ n . ( M , l 1 , t , n , h h , v ) \in E x e c \langle p r o o f \rangle \langle p r o o f \rangle
lemma MStep-Exec2:
    |(M,l,s,kb,l1,t)\inMStep; (M,l1,t,k,hh,v)\inExec\rrbracket
    \Longrightarrow \exists n . ( M , l , s , n , h h , v ) \in E x e c \langle p r o o f \rangle
```

Finally, the definition of the non-height-indexed relation.
definition $M S::$ Mbody $\Rightarrow$ Label $\Rightarrow$ State $\Rightarrow$ Label $\Rightarrow$ State $\Rightarrow$ bool
where MS M l s ll $t=(\exists k .(M, l, s, k, l l, t): M S t e p)$

### 4.2 Reachability relation

The second auxiliary operational judgement is required for the interpretation of invariants and method invariants. Invariants are expected to be satisfied in all heap components of (future) states that occur either in the same frame as the current state or a subframe thereof. Likewise, method invariants are expected to be satisfied by all heap components of states observed during the execution of a method, including subframes. None of the previous three operational judgements allows us to express these interpretations, as Step injects the execution of an invoked method as a single step. Thus, states occurring in subframes cannot be related to states occurring in the parent frame using these judgements. This motivates the introduction of predicates relating states $s$ and $t$ whenever the latter can be reach from the former, i.e. whenever $t$ occurs as a successor of $s$ in the same frame as $s$ or one of its subframes. Again, we first define a relation that includes an explicit derivation height index.

```
inductive-set
    Reachable::(Mbody }\times\mathrm{ Label }\times\mathrm{ State }\times\mathrm{ nat }\times\mathrm{ State) set
where
Reachable-zero: \llbracketk=0;t=s\rrbracket\Longrightarrow(M,l,s,k,t):Reachable
|
Reachable-step:
    \llbracket (M,l,s,n,ll,r):Step; (M,ll,r,m,t):Reachable; k=Suc m+n\rrbracket
    \Longrightarrow ( M , l , s , k , t ) : R e a c h a b l e
|
Reachable-invS:
    \llbracket mbody-is C m (par,code,l0); get-ins M l = Some (invokeS C m);
        s=(ops,S,h); (ops,par,R,ops1):Frame;
        ((par,code,l0),l0,([],R,h), n, t):Reachable; k=Suc n\rrbracket
    \Longrightarrow(M,l,s,k,t): Reachable
```

The following properties of are useful to notice.

```
\(\langle p r o o f\rangle\) lemma ZeroHeightReachableElim: \((M, l, s, 0, r) \in\) Reachable \(\Longrightarrow r=s\langle p r o o f\rangle\)
lemma ReachableSplit[rule-format]:
    \((M, l, s, k, t) \in\) Reachable \(\Longrightarrow\)
        \(1 \leq k \longrightarrow\)
    \(((\exists\) n m r ll. \((M, l, s, n, l l, r): S t e p \wedge\)
                            \((M, l l, r, m, t):\) Reachable \(\wedge\) Suc \(m+n=k) \vee\)
    ( \(\exists\) n ops \(S h\) c m par \(R\) ops1 code l0.
        \(s=(o p s, S, h) \wedge\) get-ins \(M l=\) Some \((\) invokeS \(c m) \wedge\)
        mbody-is c m (par,code,l0) ^(ops,par,R,ops1):Frame \(\wedge\)
        \(((\) par, code,\(l 0), l 0,([], R, h), n, t):\) Reachable \(\wedge S u c n=k))\langle\) proof \(\rangle\)
lemma Reachable-returnElim[rule-format]:
\((M, l, s, k, t) \in\) Reachable \(\Longrightarrow\) get-ins \(M l=\) Some vreturn \(\longrightarrow t=s\langle p r o o f\rangle\)
```

Similar to the operational semantics, we define a variation of the reachability relation that hides the index.
definition Reach::Mbody $\Rightarrow$ Label $\Rightarrow$ State $\Rightarrow$ State $\Rightarrow$ bool
where Reach Mlst=( $\exists k .(M, l, s, k, t):$ Reachable $)$

## 5 Soundness

This section contains the soundness proof of the program logic. In the first subsection, we define our notion of validity, thus formalising our intuitive explanation of the terms preconditions, specifications, and invariants. The following two subsections contain the details of the proof and can easily be skipped during a first pass through the document.

### 5.1 Validity

A judgement is valid at the program point $C$.m.l (i.e. at label $l$ in method $m$ of class $C$ ), written valid $C m l A B I$ or, in symbols,

$$
\vDash\{A\} C, m, l\{B\} I
$$

if $A$ is a precondition for $B$ and for all local annotations following $l$ in an execution of $m$, and all reachable states in the current frame or yet-to-be created subframes satisfy $I$. More precisely, whenever an execution of the method starting in an initial state $s_{0}$ reaches the label $l$ with state $s$, the following properties are implied by $A\left(s_{0}, s\right)$.

1. If the continued execution from $s$ reaches a final state $t$ (i.e. the method terminates), then that final state $t$ satisfies $B\left(s_{0}, s, t\right)$.
2. Any state $s^{\prime}$ visited in the current frame during the remaining program execution whose label carries an annotation $Q$ will satisfy $Q\left(s_{0}, s^{\prime}\right)$, even if the execution of the frame does not terminate.
3. Any state $s^{\prime}$ visited in the current frame or a subframe of the current frame will satisfy $I\left(s_{0}, s\right.$, heap $\left.\left(s^{\prime}\right)\right)$, again even if the execution does not terminate.

Formally, this interpretation is expressed as follows.

```
definition valid \(::\) Class \(\Rightarrow\) Method \(\Rightarrow\) Label \(\Rightarrow\) Assn \(\Rightarrow\) Post \(\Rightarrow\) Inv \(\Rightarrow\) bool where
valid \(C m l A B I=\)
    \((\forall\) M. mbody-is \(C m M \longrightarrow\)
    \((\forall\) Mspec Minv Anno \(\operatorname{MST} \downarrow(C, m)=\) Some \((\) Mspec, Minv,Anno \() \longrightarrow\)
    \((\forall\) par code \(l 0 . M=(\) par, code,\(l 0) \longrightarrow\)
    ( \(\forall\) s0 s. MS M l0 (mkState s0) ls \(\longrightarrow A\) s0 s \(\longrightarrow\)
    \(((\forall h v\). Opsem Mlshv\(\longrightarrow B s 0 s(h, v)) \wedge\)
        \((\forall \mathrm{ll} r .(M S M l \operatorname{ll} r \longrightarrow(\forall Q . A n n o \downarrow(l l)=S o m e Q \longrightarrow Q s 0 r)) \wedge\)
                            \((\) Reach Mls r \(\longrightarrow\) Is0 s (heap r))))))))
```

abbreviation valid-syntax :: Assn $\Rightarrow$ Class $\Rightarrow$ Method $\Rightarrow$
Label $\Rightarrow$ Post $\Rightarrow$ Inv $\Rightarrow$ bool
$(\vDash\{-\}-,-,-\{-\}-[200,200,200,200,200,200]$ 200 $)$
where valid-syntax $A C m l B I==$ valid $C m l A B I$

This notion of validity extends that of Bannwart-Müller by allowing the postcondition to differ from method specification and to refer to the initial state, and by including invariants. In the logic of Bannwart and Müller, the validity of a method specification is given by a partial correctness (Hoare-style) interpretation, while the validity of preconditions of individual instructions is such that a precondition at $l$ implies the preconditions of its immediate control flow successors.

Validity us lifted to contexts and the method specification table. In the case of the former, we simply require that all entries be valid.

## definition $G$-valid:: CTXT $\Rightarrow$ bool where

$G$-valid $G=(\forall C m l A B I . G \downarrow(C, m, l)=\operatorname{Some}(A, B, I) \longrightarrow$

$$
\vDash\{A\} C, m, l\{B\} I)
$$

Regarding the specification table, we require that the initial label of each method satisfies an assertion that ties the method precondition to the current state.
definition MST-valid ::bool where
MST-valid $=(\forall$ C m par code l0 T MI Anno.
mbody-is C m (par,code,l0) $\longrightarrow M S T \downarrow(C, m)=$ Some $(T, M I, A n n o) \longrightarrow$ $\vDash\{(\lambda s 0 \mathrm{~s} . \mathrm{s}=\mathrm{mkState} \mathrm{sO})\} C, m, l 0\{(m k$ Post $T)\}(m k$ Inv MI $))$
definition Prog-valid::bool where
Prog-valid $=(\exists G . G$-valid $G \wedge$ MST-valid $)$
The remainder of this section contains a proof of soundness, i.e. of the property

$$
V P \Longrightarrow \text { Prog-valid, }
$$

and is structured into two parts. The first step (Section 5.2) establishes a soundness result where the $V P$ property is replaced by validity assumptions regarding the method specification table and the context. In the second step (Section 5.3), we show that these validity assumptions are satisfied by verified programs, which implies the overall soundness theorem.

### 5.2 Soundness under valid contexts

The soundness proof proceeds by induction on the axiomatic semantics, based on an auxiliary lemma for method invocations that is proven by induction on the derivation height of the operational semantics. For the latter induction, relativised notions of validity are employed that restrict the derivation height of the program continuations affected by an assertion. The appropriate definitions of relativised validity for judgements, for the precondition table, and for the method specification table are as follows.

```
definition validn::
    nat \(\Rightarrow\) Class \(\Rightarrow\) Method \(\Rightarrow\) Label \(\Rightarrow\) Assn \(\Rightarrow\) Post \(\Rightarrow\) Inv \(\Rightarrow\) bool where
validn \(K C m l A B I=\)
    ( \(\forall\) M. mbody-is \(C m M \longrightarrow\)
    \((\forall\) Mspec Minv Anno \(. \operatorname{MST} \downarrow(C, m)=\) Some (Mspec,Minv,Anno) \(\longrightarrow\)
    \((\forall\) par code l0 \(. M=(\) par, code,\(l 0) \longrightarrow\)
    ( \(\forall\) so s. MS M l0 (mkState so) ls \(\mathrm{s} \longrightarrow \mathrm{AsOs} \longrightarrow\)
    \((\forall k . k \leq K \longrightarrow\)
    \(((\forall h v .(M, l, s, k, h, v): E x e c \longrightarrow B s 0 s(h, v)) \wedge\)
        ( \(\forall\) ll r \(\cdot((M, l, s, k, l l, r):\) MStep \(\longrightarrow\)
                    \((\forall Q . \operatorname{Anno} \downarrow(l l)=\) Some \(Q \longrightarrow Q\) sO \(r)) \wedge\)
                        \(((M, l, s, k, r):\) Reachable \(\longrightarrow I\) s0 \(s(\) heap \(r))))))))\) )
```

abbreviation validn-syntax :: nat $\Rightarrow$ Assn $\Rightarrow$ Class $\Rightarrow$ Method $\Rightarrow$
Label $\Rightarrow$ Post $\Rightarrow$ Inv $\Rightarrow$ bool
$(\models-\{-\}-,-,-\{-\}-[200,200,200,200,200,200,200] 200)$
where validn-syntax $K A C m l B I==\operatorname{validn} K C m l A B I$
definition $G$-validn::nat $\Rightarrow C T X T \Rightarrow$ bool where
$G$-validn $K G=(\forall C m l A B I . G \downarrow(C, m, l)=\operatorname{Some}(A, B, I) \longrightarrow$
$\left.\models_{K}\{A\} C, m, l\{B\} I\right)$
definition MST-validn $::$ nat $\Rightarrow$ bool where
MST-validn $K=(\forall C$ m par code $l 0 T M I$ Anno.
mbody-is $C m$ (par,code,$l 0) \longrightarrow M S T \downarrow(C, m)=\operatorname{Some}(T, M I, A n n o) \longrightarrow$
$\models_{K}\{(\lambda s 0 \mathrm{~s} . \mathrm{s}=\mathrm{mkState} \mathrm{s} 0)\} C, m, l 0\{(m k$ Post $T)\}(m k$ Inv MI) $)$
definition Prog-validn::nat $\Rightarrow$ bool where
Prog-validn $K=(\exists G . G$-validn $K G \wedge$ MST-validn $K)$
The relativised notions are related to each other, and to the native notions of validity as follows.

```
lemma valid-validn: \(\models\{A\} C, m, l\{B\} I \Longrightarrow \models_{K}\{A\} C, m, l\{B\} I\langle\) proof \(\rangle\)
```

lemma validn-valid: $\llbracket \forall K . \models_{K}\{A\} C, m, l\{B\} I \rrbracket \Longrightarrow \vDash\{A\} C, m, l\{B\}$
$I\langle$ proof $\rangle$
lemma validn-lower:
$\llbracket \mid={ }_{K}\{A\} C, m, l\{B\} I ; L \leq K \rrbracket \Longrightarrow \models_{L}\{A\} C, m, l\{B\} I\langle$ proof $\rangle$
lemma $G$-valid-validn: $G$-valid $G \Longrightarrow G$-validn $K G\langle$ proof $\rangle$
lemma G-validn-valid: $\llbracket \forall K . G$-validn $K G \rrbracket \Longrightarrow G$-valid $G\langle p r o o f\rangle$
lemma $G$-validn-lower: $\llbracket G$-validn $K G ; L \leq K \rrbracket \Longrightarrow G$-validn $L G\langle$ proof $\rangle$
lemma MST-validn-valid: $\llbracket \forall K . M S T$-validn $K \rrbracket \Longrightarrow M S T$-valid $\langle p r o o f\rangle$
lemma MST-valid-validn:MST-valid $\Longrightarrow$ MST-validn $K\langle p r o o f\rangle$
lemma MST-validn-lower: $\llbracket M S T$-validn $K ; L \leq K \rrbracket \Longrightarrow M S T$-validn $L\langle p r o o f\rangle$

We define an abbreviation for the side conditions of the rule for static method invocations...

```
definition \(I N V S-S C:\) :
    Class \(\Rightarrow\) Method \(\Rightarrow\) Label \(\Rightarrow\) Class \(\Rightarrow\) Method \(\Rightarrow\) MethSpec \(\Rightarrow\) MethInv \(\Rightarrow\)
    ANNO \(\Rightarrow\) ANNO \(\Rightarrow\) Mbody \(\Rightarrow\) Assn \(\Rightarrow\) Inv \(\Rightarrow\) bool where
INVS-SC C m l D m \({ }^{\prime}\) T MI Anno Anno2 \(M^{\prime} A I=(\exists\) M par code l0 T1 MI1.
    mbody-is \(C m M \wedge\) get-ins \(M l=\) Some (invokeS \(\left.D m^{\prime}\right) \wedge\)
    \(M S T \downarrow(C, m)=S o m e(T 1, M I 1, A n n o) \wedge\)
    \(M S T \downarrow\left(D, m^{\prime}\right)=\) Some \((T, M I, A n n o 2) \wedge\)
    mbody-is \(D m^{\prime} M^{\prime} \wedge M^{\prime}=(\) par, code,\(l 0) \wedge\)
    \((\forall Q . A n n o \downarrow(l)=\) Some \(Q \longrightarrow(\forall s 0 s . A s 0 s \longrightarrow Q s 0 s)) \wedge\)
    \((\forall s 0 s . A s 0 s \longrightarrow I s 0 s(\) heap \(s)) \wedge\)
    ( \(\forall\) s0 ops1 ops2 \(S R h t\). (ops1,par, R,ops2) : Frame \(\longrightarrow\)
            As0 \((o p s 1, S, h) \longrightarrow M I(R, h) t \longrightarrow I s 0(o p s 1, S, h) t))\)
```

... and another abbreviation for the soundness property of the same rule.

```
definition INVS-soundK::
    nat }=>\mathrm{ CTXT }=>\mathrm{ Class }=>\mathrm{ Method }=>\mathrm{ Label }=>\mathrm{ Class }=>\mathrm{ Method }
    MethSpec }=>\mathrm{ MethInv }=>\mathrm{ ANNO }=>\mathrm{ ANNO }=>\mathrm{ Mbody }=>\mathrm{ Assn }
    Post }=>\mathrm{ Inv }=>\mathrm{ bool where
INVS-soundK K GCml D m' T MI Anno Anno2 M'A B I =
    (INVS-SC C m l D m' T MI Anno Anno2 M'A I \longrightarrow
    G-validn K G \longrightarrow MST-validn K}
    \models
        {(SINV-post (fst M')TB)} (SINV-inv (fst M')T I)
    \longrightarrow \models ( K + 1 ) \{ A \} C , m , l \{ B \} I )
```

The proof that this property holds for all $K$ proceeds by induction on $K$.

```
lemma INVS-soundK-all:
    INVS-soundK KGCml D m' T MI Anno Anno2 M' A B I\langleproof\rangle
```

The heart of the soundness proof - the induction on the axiomatic semantics.

```
lemma SOUND-Aux[rule-format]:
(b,G,C,m,l,A,B,I):SP-Judgement \Longrightarrow G-validn K G MST-validn K }
```



```
    ((\negb)\longrightarrow\models(Suc K) {A}C,m,l{{B}I))\langleproof\rangle
```

The statement of this lemma gives a semantic interpretation of the two judgement forms, as $S P$-Assum-judgements enjoy validity up to execution height $K$, while $S P$-Deriv-judgements are valid up to level $K+1$.
$\langle p r o o f\rangle$
From this, we obtain a soundness result that still involves context validity.

## theorem SOUND-in-CTXT:

$\llbracket G \triangleright\{A\} C, m, l\{B\} I ; G$-valid $G ; M S T$-valid $\rrbracket \Longrightarrow \vDash\{A\} C, m, l\{B\} I\langle$ proof $\rangle$
We will now show that the two semantic assumptions can be replaced by the verified-program property.

### 5.3 Soundness of verified programs

In order to obtain a soundness result that does not require validity assumptions of the context or the specification table, we show that the $V P$ property implies context validity. First, the elimination of contexts. By induction on $k$ we prove
lemma VPG-MSTn-Gn[rule-format]:
VP-G $G \longrightarrow$ MST-validn $k \longrightarrow G$-validn $k G\langle$ proof $\rangle$
which implies
lemma VPG-MST-G: $\llbracket V P-G ~ G ; ~ M S T-v a l i d \rrbracket \Longrightarrow G$-valid $G\langle p r o o f\rangle$
Next, the elimination of MST-valid. Again by induction on $k$, we prove lemma $V P G$-MSTn[rule-format]: VP-G $G \longrightarrow$ MST-validn $k\langle$ proof $\rangle$
which yields
lemma $V P G$-MST:VP-G $G \Longrightarrow$ MST-valid $\langle$ proof $\rangle$
Combining these two results, and unfolding the definition of program validity yields the final soundness result.
theorem VP-VALID: VP $\Longrightarrow$ Prog-valid $\langle$ proof $\rangle\langle$ proof $\rangle$

## 6 A derived logic for a strong type system

In this section we consider a system of derived assertions, for a type system for bounded heap consumption. The type system arises by reformulating the analysis of Cachera, Jensen, Pichardie, and Schneider [4] for a highlevel functional language. The original approach of Cachera et al. consists of formalising the correctness proof of a certain analysis technique in Coq. Consequently, the verification of a program requires the execution of the analysis algorithm inside the theorem prover, which involves the computation of the (method) call graph and fixed point iterations. In contrast, our
approach follows the proof-carrying code paradigm more closely: the analysis amounts to a type inference which is left unformalised and can thus be carried out outside the trusted code base. Only the result of the analysis is communicated to the code recipient. The recipient verifies the validity of the certificate by a largely syntax-directed single-pass traversal of the (low-level) code using a domain-specific program logic. This approach to proof-carrying code was already explored in the MRG project, with respect to program logics of partial correctness [3] and a type system for memory consumption by Hofmann and Jost [7]. In order to obtain syntax-directedness of the proof rules, these had to be formulated at the granularity of typing judgements. In contrast, the present proof system admits proof rules for individual JVM instructions.
Having derived proof rules for individual JVM instructions, we introduce a type system for a small functional language, and a compilation into bytecode. The type system associates a natural number $n$ to an expression $e$, in a typing context $\Sigma$. Informally, the interpretation of a typing judgement $\Sigma \triangleright e: n$ is that the evaluation of $e$ (which may include the invocation of functions whose resource behaviour is specified in $\Sigma$ ) does not perform more than $n$ allocations. The type system is then formally proven sound, using the derived logic for bytecode. By virtue of the invariants, the guarantee given by the present system is stronger than the one given by our encoding of the Hofmann-Jost system, as even non-terminating programs can be verified in a meaningful way.

### 6.1 Syntax and semantics of judgements

The formal interpretation at JVM level of a type $n$ is given by a triple

$$
\operatorname{Cachera}(n)=(A, B, I)
$$

consisting of a (trivial) precondition, a post-condition, and a strong invariant.
definition Cachera::nat $\Rightarrow$ (Assn $\times$ Post $\times$ Inv $)$ where
Cachera $n=(\lambda s 0$ s. True,

$$
\begin{aligned}
& \lambda s 0(o p s, s, h)(k, v) \cdot|k| \leq|h|+n, \\
& \lambda s 0(o p s, s, h) k .|k| \leq|h|+n)
\end{aligned}
$$

This definition is motivated by the expectation that $\triangleright\{A\}\ulcorner e\urcorner\{B\} I$ should be derivable whenever the type judgement $\Sigma \triangleright e: n$ holds, where $\ulcorner e\urcorner$ is the translation of compiling the expression $e$ into JVML, and the specification table $M S T$ contains the interpretations of the entries in $\Sigma$.

We abbreviate the above construction of judgements by a predicate deriv.
definition deriv::CTXT $\Rightarrow$ Class $\Rightarrow$ Method $\Rightarrow$ Label $\Rightarrow$ (Assn $\times$ Post $\times$ Inv $) \Rightarrow$ bool where

```
deriv \(G C m l(A B I)=(\operatorname{let}(A, B, I)=A B I \operatorname{in}(G \triangleright\{A\} C, m, l\{B\} I))\)
```

Thus，the intended interpretation of a typing judgement $\Sigma \triangleright e: n$ is deriv $C m l($ Cachera $n)$
if $e$ translates to a code block whose first instruction is at C．m．l．
We also define a judgement of the auxiliary form of sequents．

```
definition derivAssum::CTXT \(\Rightarrow\) Class \(\Rightarrow\) Method \(\Rightarrow\) Label \(\Rightarrow\)
    (Assn \(\times\) Post \(\times\) Inv \() \Rightarrow\) bool where
derivAssum \(G C m l(A B I)=(\) let \((A, B, I)=A B I\) in \(G \triangleright\langle A\rangle C, m, l\langle B\rangle I)\)
```

The following operation converts a derived judgement into the syntactical form of method specifications．

```
definition \(m k S P E C::(\) Assn \(\times\) Post \(\times I n v) \Rightarrow A N N O \Rightarrow\)
            \((\) MethSpec \(\times\) MethInv \(\times A N N O)\) where
\(m k S P E C(A B I)\) Anno \(=(\) let \((A, B, I)=A B I\) in
    ( \(\lambda s 0 t . B s 0(m k S t a t e ~ s 0) t, \lambda s 0 h . I s 0(m k S t a t e ~ s 0) h, A n n o))\)
```

This enables the interpretation of typing contexts $\Sigma$ as a set of constraints on the specification table MST．

## 6．2 Derived proof rules

We are now ready to prove derived rules，i．e．proof rules where assumptions as well as conclusions are of the restricted assertion form．While their jus－ tification unfolds the definition of the predicate deriv，their application will not．We first give syntax－directed proof rules for all JVM instructions：

```
lemma \(C A C H-N E W\) :
    【ins-is C ml (new c); MST \(\downarrow(C, m)=\) Some(Mspec,Minv,Anno);
        Anno \(\downarrow(l)=\) None \(; n=k+1 ;\) derivAssum \(G C m(l+1)(\) Cachera \(k) \rrbracket\)
    \(\Longrightarrow\) deriv \(G C m l(\) Cachera \(n)\langle\) proof \(\rangle\)
lemma CACH-INSTR:
    【ins-is Cmli;
        \(I \in\{\) const \(c\), dup, pop, swap, load \(x\), store \(x\), binop \(f\),
                unop \(g\), getfield \(d F\), putfield \(d F\), checkcast \(d\}\);
        \(\operatorname{MST} \downarrow(C, m)=\operatorname{Some}(M s p e c, M i n v, A n n o) ;\) Anno \(\downarrow(l)=\) None;
        derivAssum G C m (l+1) (Cachera n) 】
    \(\Longrightarrow\) deriv \(G C m l(\) Cachera \(n)\langle\) proof \(\rangle\)
lemma CACH-RET:
    【ins-is \(C\) m l vreturn \(; \operatorname{MST} \downarrow(C, m)=\operatorname{Some}(M s p e c, M i n v, A n n o)\);
        Anno \(\downarrow(l)=\) None 】
    \(\Longrightarrow\) deriv \(G\) C ml (Cachera 0) \(\langle\) proof \(\rangle\)
lemma CACH-GOTO:
    【ins-is \(C m l(\) goto pc); MST \(\downarrow(C, m)=\operatorname{Some}(M s p e c, M i n v, A n n o)\);
        Anno \(\downarrow(l)=\) None; derivAssum \(G C m p c(\) Cachera \(n) \rrbracket\)
    \(\Longrightarrow\) deriv \(G C m l(\) Cachera \(n)\langle\) proof \(\rangle\)
```

```
lemma CACH-IF:
    |ins-is Cml (iftrue pc);MST\downarrow (C,m)=Some(Mspec,Minv,Anno);
        Anno\downarrow(l)=None; derivAssum G C m pc (Cachera n);
        derivAssum G Cm (l+1) (Cachera n)】
    #deriv G Cml (Cachera n)\langleproof\rangle
lemma CACH-INVS:
    \llbracket ins-is Cml (invokeS D m'); mbody-is D m' (par,code,l0);
        MST\downarrow (C,m)=Some(Mspec,Minv,Anno);Anno\downarrow }(l)=None
        MST\downarrow(D, m') = Some(mkSPEC (Cachera k) Anno2);
        nk=n+k; derivAssum GC m (l+1) (Cachera n)\rrbracket
    "deriv G C m l (Cachera nk)\langleproof\rangle
```

In addition, we have two rules for subtyping

```
lemma \(C A C H-S U B\) :
    \(\llbracket\) deriv \(G C m l(\) Cachera \(n) ; n \leq k \rrbracket \Longrightarrow\) deriv \(G C m l(\) Cachera \(k)\langle p r o o f\rangle\)
lemma CACHAssum-SUB:
    \(\llbracket\) derivAssum \(G C m l(\) Cachera \(n) ; n \leq k \rrbracket\)
    \(\Longrightarrow\) derivAssum \(G C m l(\) Cachera \(k)\langle\) proof \(\rangle\)
```

and specialised forms of the axiom rule and the injection rule.

```
lemma \(C A C H-A X\) :
    \(\llbracket G \downarrow(C, m, l)=\) Some \((\) Cachera \(n) ; M S T \downarrow(C, m)=\operatorname{Some}(M s p e c, M i n v, A n n o) ;\)
        Anno \(\downarrow(l)=N o n e \rrbracket\)
    \(\Longrightarrow\) derivAssum \(G C m l(\) Cachera \(n)\langle\) proof \(\rangle\)
lemma \(C A C H-I N J E C T\) :
    deriv \(G C m l(\) Cachera \(n) \Longrightarrow\) derivAssum \(G C m l(\) Cachera \(n)\langle\) proof \(\rangle\)
```

Finally, a verified-program rule relates specifications to judgements for the method bodies. Thus, even the method specifications may be given as derived assertions (modulo the $m k S P E C$-conversion).

```
lemma \(\mathrm{CACH}-\mathrm{VP}\) :
\(\llbracket \forall\) c m par code l0. mbody-is c m (par, code, l0) \(\longrightarrow\)
        ( \(\exists\) n Anno . \(\operatorname{MST} \downarrow(c, m)=\operatorname{Some}(m k S P E C(\) Cachera \(n)\) Anno \() \wedge\)
                        deriv G c m l0 (Cachera n));
        \(\forall c m l A B I . G \downarrow(c, m, l)=\operatorname{Some}(A, B, I) \longrightarrow\)
                            \((\exists n \cdot(A, B, I)=\) Cachera \(n \wedge\) deriv \(G c m l(\) Cachera \(n)) \rrbracket\)
\(\Longrightarrow V P\langle\) proof \(\rangle\)
```


### 6.3 Soundness of high-level type system

We define a first-order functional language where expressions are stratified into primitive expressions and general expressions. The language supports the construction of lists using constructors NilPrim and ConsPrim ht, and includes a corresponding pattern match operation. In order to simplify the compilation, function identifiers are taken to be pairs of class names and method names.
type-synonym Fun $=$ Class $\times$ Method

```
datatype Prim \(=\)
    IntPrim int
| UnPrim Val \(\Rightarrow\) Val Var
| BinPrim Val \(\Rightarrow\) Val \(\Rightarrow\) Val Var Var
| NilPrim
| ConsPrim Var Var
| CallPrim Fun Var list
datatype Expr =
    PrimE Prim
| LetE Var Prim Expr
| CondE Var Expr Expr
| MatchE Var Expr Var Var Expr
```

type-synonym FunProg $=($ Fun, Var list $\times$ Expr $)$ AssList

The type system uses contexts that associate a type (natural number) to function identifiers.
type-synonym TP-Sig $=($ Fun, nat $)$ AssList
We first give the rules for primitive expressions.

```
inductive-set TP-prim::(TP-Sig }\times\mathrm{ Prim }\times\mathrm{ nat)set
```

where
TP-int: ( $\Sigma$, IntPrim i,0) : TP-prim
|
TP-un: $(\Sigma, U n P r i m f x, 0):$ TP-prim
|
TP-bin: $(\Sigma$, BinPrim $f x y, 0)$ : TP-prim
|
TP-nil: $(\Sigma$, NilPrim, 0$): T P-p r i m$
TP-cons: $(\Sigma$, ConsPrim $x y, 1)$ : TP-prim
|
TP-Call: $\llbracket \Sigma \downarrow f=$ Some $n \rrbracket \Longrightarrow(\Sigma$, CallPrim $f$ args,$n):$ TP-prim

Next, the rules for general expressions.

```
inductive-set TP-expr::(TP-Sig }\times\mathrm{ Expr }\times\mathrm{ nat) set
where
TP-sub:\llbracket(\Sigma,e,m):TP-expr;m\leqn\rrbracket\Longrightarrow (\Sigma,e,n):TP-expr
|
TP-prim:\llbracket(\Sigma,p,n):TP-prim\rrbracket\Longrightarrow(\Sigma,PrimE p,n):TP-expr
|
TP-let:\llbracket(\Sigma,p,k):TP-prim; (\Sigma,e,m):TP-expr; n =k+m\rrbracket
    \Longrightarrow(\Sigma,LetE x pe,n):TP-expr
|
TP-Cond:\llbracket(\Sigma,e1,n):TP-expr; (\Sigma,e2,n):TP-expr\rrbracket
    \Longrightarrow ( \Sigma , C o n d E ~ x ~ e 1 ~ e 2 , n ) ~ : ~ T P - e x p r ~
```

```
TP-Match:\llbracket(\Sigma,e1,n):TP-expr; (\Sigma,e2,n):TP-expr\rrbracket
    \Longrightarrow(\Sigma,MatchE x e1 ht e2,n):TP-expr
```

A functional program is well-typed if its domain agrees with that of some context such that each function body validates the context entry.

```
definition TP::TP-Sig => FunProg }=>\mathrm{ bool where
TP \SigmaF=((\forallf. (\Sigma\downarrowf=None) = (F\downarrowf=None))}
    (\forall f n par e. }\Sigma\downarrowf=Some n\longrightarrowF\downarrowf=Some (par,e)\longrightarrow(\Sigma,e,n):TP-expr)
```

For the translation into bytecode, we introduce identifiers for a class of lists, the expected field names, and a temporary (reserved) variable name.

```
axiomatization
    LIST::Class and
    HD::Field and
    TL::Field and
    tmp::Var
```

The compilation of primitive expressions extends a code block by a sequence of JVM instructions that leave a value on the top of the operand stack.

```
inductive-set compilePrim::
    (Label }\times(\mathrm{ Label,Instr ) AssList }\times\mathrm{ Prim }\times((\mathrm{ Label,Instr ) AssList }\times\mathrm{ Label ) ) set
where
compileInt: (l, code, IntPrim i, (code[l\mapsto(const (IVal i))],l+1)) : compilePrim
|
compileUn:
    (l, code, UnPrim f x, (code[l\mapsto(load x)][(l+1)\mapsto(unop f)],l+2)) : compilePrim
|
compileBin:
    (l, code, BinPrim f x y,
        (code[l\mapsto(load x)][(l+1)\mapsto(load y)][(l+2)\mapsto(binop f)],l+3)) : compilePrim
|
compileNil:
    (l, code, NilPrim, (code[l\mapsto(const (RVal Nullref))],l+1)) : compilePrim
|
compileCons:
    (l, code, ConsPrim x y,
        (code[l\mapsto(load y)][(l+1)\mapsto(load x)]
            [(l+2)\mapsto(new LIST)][(l+3)\mapstostore tmp]
            [(l+4)\mapstoload tmp][(l+5)\mapsto(putfield LIST HD)]
            [(l+6)\mapstoload tmp][(l+7)\mapsto(putfield LIST TL)]
            [(l+8)\mapsto(load tmp)],l+9)) : compilePrim
|
compileCall-Nil:
    (l, code, CallPrim f [],(code[l\mapstoinvokeS (fst f) (snd f)],l+1)): compilePrim
|
compileCall-Cons:
    \llbracket(l+1,code[l\mapstoload x], CallPrim f args,OUT) : compilePrim\rrbracket
    \Longrightarrow ( l , ~ c o d e , ~ C a l l P r i m ~ f ~ ( x \# a r g s ) , ~ O U T ) : ~ c o m p i l e P r i m ~
```

The following lemma shows that the resulting code is an extension of the code submitted as an argument, and that the new instructions define a contiguous block.

```
lemma compilePrim-Prop 1 [rule-format]:
( \(l\), code, \(p\), OUT) : compilePrim \(\Longrightarrow\)
    \((\forall\) code1l1. OUT \(=(\) code1,\(l 1) \longrightarrow\)
    \((l<l 1 \wedge(\forall l l . l l<l \longrightarrow\) code \(1 \downarrow l l=\operatorname{code} \downarrow l l) \wedge\)
    \((\forall l l . l \leq l l \longrightarrow l l<l 1 \longrightarrow(\exists\) ins \(. \operatorname{code} 1 \downarrow l l=\) Some ins \())))\langle\) proof \(\rangle\)
```

A signature corresponds to a method specification table if all context entries are represented as $M S T$ entries and method names that are defined in the global program $P$.
definition Sig-good::TP-Sig $\Rightarrow$ bool where
Sig-good $\Sigma=$
$(\forall C m n . \Sigma \downarrow(C, m)=$ Some $n \longrightarrow$
$(\operatorname{MST} \downarrow(C, m)=$ Some $(m k S P E C($ Cachera $n)$ emp $) \wedge$
( $\exists$ par code l0 . mbody-is C m (par, code,l0))))
This definition requires $M S T$ to associate the specification

$$
m k S P E C(\text { Cachera } n) \text { emp }
$$

to each method to which the type signature associates the type $n$. In particular, this requires the annotation table of such a method to be empty. Additionally, the global program $P$ is required to contain a method definition for each method (i.e. function) name occurring in the domain of the signature.

An auxiliary abbreviation that captures when a block of code has trivial annotations and only comprises defined program labels.

```
definition Segment::
    Class \(\Rightarrow\) Method \(\Rightarrow\) Label \(\Rightarrow\) Label \(\Rightarrow\) (Label,Instr)AssList \(\Rightarrow\) bool
where
Segment C m l l1 code \(=\)
    ( \(\exists\) Mspec Minv Anno \(. \operatorname{MST} \downarrow(C, m)=\) Some (Mspec, Minv,Anno) \(\wedge\)
        \((\forall l l . l \leq l l \longrightarrow l l<l 1 \longrightarrow\)
            Anno \(\downarrow(l l)=\) None \(\wedge(\exists\) ins. ins-is \(C m l l\) ins \(\wedge\) code \(\downarrow l l=\) Some ins \()))\)
\(\langle p r o o f\rangle\langle p r o o f\rangle\langle p r o o f\rangle\langle p r o o f\rangle\)
```

The soundness of (the translation of) a function call is proven by induction on the list of arguments.

```
lemma Call-SoundAux[rule-format]:
    \Sigma\downarrowf=Some n \longrightarrow
        MST\downarrow (fst f,snd f) = Some(mkSPEC (Cachera n) Anno2) }
        (\exists par body l0 . mbody-is (fst f) (snd f) (par,body,l0))\longrightarrow
        ( }\foralll\mathrm{ code code1 l1 G C m T MI k.
            (l, code, CallPrim f args, code1,l1) \in compilePrim }
```

```
MST\downarrow (C,m) = Some (T,MI,Anno) \longrightarrowSegment C m l l1 code1 \longrightarrow
derivAssum G Cml1 (Cachera k)\longrightarrow
deriv GCml (Cachera ( }n+k)))\langle\mathrm{ proof }
lemma Call-Sound:
    | Sig-good \Sigma; \Sigma\downarrowf=Some n;
    (l, code, CallPrim f args, code1, l1) \in compilePrim;
    MST\downarrow (C,m) = Some (T, MI,Anno); Segment C m l l1 code1;
    derivAssum G Cml1 (Cachera nn); k=n+nn\rrbracket
deriv GCml (Cachera k)\langleproof\rangle
```

The definition of basic instructions.

```
definition basic::Instr \(\Rightarrow\) bool where
basic ins \(=((\exists c\). ins \(=\) const \(c) \vee\) ins \(=d u p \vee\)
    ins \(=\) pop \(\vee\) ins \(=\) swap \(\vee(\exists x\). ins \(=\) load \(x) \vee\)
    \((\exists y\). ins \(=\) store \(y) \vee(\exists f\). ins \(=\) binop \(f) \vee\)
    \((\exists \mathrm{g}\). ins \(=\) unop \(g) \vee(\exists\) c1 F1. ins \(=\) getfield \(c 1 F 1) \vee\)
    ( \(\exists\) c2 F2. ins \(=\) putfield c2 F2) \(\vee\)
    \((\exists\) c3. ins \(=\) checkcast \(c 3)\) )
```

Next, we prove the soundness of basic instructions. The hypothesis refers to instructions located at the program continuation.

## lemma Basic-Sound:

$\llbracket$ Segment $C$ m l ll code; $\operatorname{MST} \downarrow(C, m)=$ Some (T, MI,Anno); l $l \leq l 1 ; l 1<l l$; $l 2=l 1+1 ;$ code $\downarrow l 1=$ Some ins; basic ins; derivAssum G C m l2 (Cachera n) $\rrbracket$
$\Longrightarrow$ deriv $G C m l 1$ (Cachera $n$ ) $\langle$ proof $\rangle$
Following this, the soundness of the type system for primitive expressions. The proof proceeds by induction on the typing judgement.

```
lemma TP-prim-Sound[rule-format]:
    (\Sigma,p,n):TP-prim \Longrightarrow
    Sig-good \Sigma\longrightarrow
( }\forall\mathrm{ l code code1 l1 G C m T MI Anno nn k.
    (l, code, p,(code1,l1)) : compilePrim }
    MST\downarrow (C,m) = Some (T,MI,Anno) }
    Segment C m l l1 code1 \longrightarrow derivAssum G C m l1 (Cachera nn) \longrightarrow
    k=n+nn\longrightarrowderiv GCml (Cachera k))\langleproof\rangle
```

The translation of general expressions is defined similarly, but no code continuation is required.

```
inductive-set compileExpr::
    (Label }\times(\mathrm{ Label,Instr ) AssList }\times\mathrm{ Expr }\times((\mathrm{ Label,Instr ) AssList }\times\mathrm{ Label ) ) set
where
compilePrimE:
    \llbracket(l, code, p,(code1,l1)) : compilePrim; OUT = (code1[l1\mapstovreturn],l1+1)\rrbracket
        \Longrightarrow(l, code, PrimE p,OUT):compileExpr
|
compileLetE:
    \llbracket(l, code, p,(code1,l1)): compilePrim; (code2,l2) = (code1[l1\mapsto(store x)],l1+1);
```

```
        (l2, code2, e, OUT) : compileExpr\rrbracket
    \Longrightarrow ( l , ~ c o d e , L e t E ~ x ~ p ~ e , O U T ) ~ : ~ c o m p i l e E x p r ~
|
compileCondE:
    \llbracket(l+2, code, e2, (codeElse,XXX)) : compileExpr;
        (XXX, codeElse, e1, (codeThen,YYY) ) : compileExpr ;
        OUT = (codeThen[l\mapstoload x][(l+1)\mapsto(iftrue XXX)], YYY)\rrbracket
    (l, code, CondE x e1 e2, OUT): compileExpr
|
compileMatchE:
    \llbracket(l+9, code, e2, (codeCons,lNil)) : compileExpr;
        (lNil, codeCons, e1, (codeNil,lRes) ) : compileExpr ;
        OUT = (codeNil[l\mapsto(load x)]
                                    [(l+1)\mapsto(unop (\lambdav. if v=RVal Nullref
                                    then TRUE else FALSE))]
                                    [(l+2)\mapsto(iftrue lNil)]
                                    [(l+3)\mapsto(load x)]
                                    [(l+4)\mapsto(getfield LIST HD)]
                                    [(l+5)\mapsto(store h)]
                                    [(l+6)\mapsto(load x)]
                                    [(l+7)\mapsto(getfield LIST TL)]
                                    [(l+8)\mapsto(store t)],lRes)]
    \Longrightarrow(l, code,MatchE x e1 h t e2,OUT): compileExpr
```

Again, we prove an auxiliary result on the emitted code, by induction on the compilation judgement.

```
lemma compileExpr-Prop1 [rule-format]:
(l,code,e,OUT) : compileExpr \(\Longrightarrow\)
    \((\forall\) code1 l1 . OUT \(=(\) code1,\(l 1) \longrightarrow\)
        \((l<l 1 \wedge\)
            \((\forall l l . l l<l \longrightarrow\) code \(1 \downarrow l l=\) code \(\downarrow l l) \wedge\)
            \((\forall l l . l \leq l l \longrightarrow l l<l 1 \longrightarrow(\exists\) ins \(. \operatorname{code} 1 \downarrow l l=\) Some ins \())))\langle\) proof \(\rangle\)
```

Then, soundness of the epxression type system is proven by induction on the typing judgement.

```
lemma TP-epxr-Sound[rule-format]:
(\Sigma,e,n):TP-expr \Longrightarrow Sig-good \Sigma\longrightarrow
( }\forall\mathrm{ l code code1 l1 G C m T MI Anno.
    (l, code, e,(code1,l1)):compileExpr \longrightarrow
    MST\downarrow (C,m) = Some (T,MI,Anno) \longrightarrow
    Segment C m l l1 code1 \longrightarrow deriv G C m l (Cachera n))\langleproof\rangle
```

The full translation of a functional program into a bytecode program is defined as follows.
definition compileProg::FunProg $\Rightarrow$ bool where
compileProg $F=$
$((\forall C$ m par e. $F \downarrow(C, m)=\operatorname{Some}($ par,$e) \longrightarrow$
( $\exists$ code l0 l. mbody-is $C m$ (rev par,code,l0) $\wedge$

$$
\begin{gathered}
(l 0,[], e,(\text { code }, l)): \text { compileExpr })) \wedge \\
(\forall C m \cdot(\exists M \cdot \operatorname{mbody-is~} m M)=(\exists \text { fdecl } \cdot F \downarrow(C, m)=\text { Some fdecl })))
\end{gathered}
$$

The final condition relating a typing context to a method specification table.

```
definition TP-MST::TP-Sig \(\Rightarrow\) bool where
TP-MST \(\Sigma=\)
    \((\forall C m\). case \((M S T \downarrow(C, m))\) of
        None \(\Rightarrow \Sigma \downarrow(C, m)=\) None
        \(\mid \operatorname{Some}(T, M I, A n n o) \Rightarrow A n n o=e m p \wedge\)
        \((\exists n . \Sigma \downarrow(C, m)=\) Some \(n \wedge\)
        \((T, M I, A n n o)=m k S P E C(\) Cachera \(n) e m p))\)
```

For well-typed programs, this property implies the earlier condition on signatures.
lemma translation-good: $\llbracket$ compileProg $F ; T P-M S T \Sigma ; T P \Sigma F \rrbracket \Longrightarrow$ Sig-good $\Sigma\langle$ proof $\rangle$

We can thus prove that well-typed function bodies satisfy the specifications asserted by the typing context.

```
lemma \(C A C H\)-BodiesDerivable[rule-format]:
    【 mbody-is C m (par, code, l); compileProg F; TP-MST \(\Sigma ; T P \Sigma F \rrbracket\)
    \(\Longrightarrow \exists n \cdot \operatorname{MST} \downarrow(C, m)=\operatorname{Some}(m k S P E C(\) Cachera \(n) e m p) \wedge\)
        deriv [] \(C m l(\) Cachera \(n)\langle\) proof \(\rangle\)
```

From this, the overall soundness result follows easily.
theorem CACH-VERIFIED: $\llbracket T P \Sigma F ; T P-M S T \Sigma$; compileProg $F \rrbracket \Longrightarrow V P\langle p r o o f\rangle$

## References

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[^0]:    ${ }^{1}$ In the paper, we slightly abstract from this by including a successor functions on labels

