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ARITHMETICAL COMPLETENESS IN LOGICS OF PROGRAMS

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Arithmetical Completeness in Logics of Programs

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Abstract.

We consider the problem of designing arithmetically complete axiom systems for proving general properties of programs; i.e., axiom systems which are complete over arithmetical universes, when all first-order formulae which are valid in such universes are taken as axioms. We prove a general Theorem of Completeness which takes care of a major part of the responsibility when designing such systems. It is then shown that what is left to do in order to establish an arithmetical completeness result, such as those appearing in [12] and [14] for the logics DL and DL+, can be described as a chain of reasoning which involves some simple utilizations of arithmetical induction. An immediate application of these observations is given in the form of an arithmetical completeness result for a new logic similar to that of Salwicki [22]. Finally, we contrast this discipline with Cook's [5] notion of relative completeness.

1. Introduction.

In the past ten years substantial progress has been made in the field of logics of programs, a field which concerns itself with the design of mathematical tools (formal and informal) for reasoning about programs. Such logics should provide means for expressing interesting properties of programs, and, hopefully, methods for proving them when true. In this paper we would like to focus on proof theory and to single out a specific trend in its development.

The logics we consider are the formal logics DL (dynamic logic) and DL+. DL was suggested as a powerful logic of programs by Pratt [21] and was further developed in [12] and [9]. DL+ was defined in [14]. These logics are extensions of first-order predicate calculus, and are capable of expressing a very wide variety of properties such as partial and total correctness, equivalence of programs, etc.

We are interested in the formulae of DL (DL-wffs) which are true in all states of a given universe U (U-valid DL-wffs). With each universe U there is associated a domain D, and in all states of U, function and predicate symbols are interpreted as functions and predicates over D. Our approach is to consider those universes A (arithmetical universes)
in which the domain essentially includes the natural numbers, and in which +, * and 0
are given their standard interpretation. In this framework one can show that with each
DL-wff P there is associated a first-order formula F_p which is equivalent to P in any
state in A.

Our approach to proving the fact that a DL-wff P is A-valid is to design an axiom
system with which P can be syntactically transformed into F_p. Then, for any arithmetical
universe A, all A-valid first-order formulae are taken as axioms, and thus, once this
transformation has been carried out, the set of axioms should contain F_p iff P is
A-valid, and thus the proof is complete. This approach we call arithmetical
axiomatization. If the axiom system has the property that any A-valid DL-wff P can indeed
be transformed into the A-equivalent first-order F_p, we say that it is an arithmetically
complete axiom system, for then a proof of it is guaranteed to exist.

Section 2 contains rigorous definitions of the concepts of state and universe and
of the logics DL and DL^+. In Section 3 we define the notion of an arithmetical universe,
present axiom systems for DL and DL^+, and state the arithmetical completeness results we
were able to prove for them in [12] and [14].

Sections 4 and 5, the main sections of the paper, are devoted to the establishment
of a general framework in which the process of designing such axiom systems and proving
their arithmetical completeness can be derived out of simple inductive considerations. In
Section 4 we prove a very general theorem, the Theorem of Completeness, in which the main
inductive step in the proofs of all the completeness results appearing in [12], [14], [10]
and [9] is captured in an abstract way. Thus, it can be seen that for any "DL-like"
logic, once arithmetical completeness can be established for very simple formulae which
involve only one appearance of a program this theorem can be applied to obtain the
completeness result for the full logic. In Section 5 we describe a way in which these
"basic" results for one-program formulae can be developed for DL and DL^+ using
straightforward reasoning. Thus, the intuition behind the inductive assertion method of
Floyd [7] and Hoare [15], which is embodied in our axiomatization of DL, can be viewed as
a special case of a rather broad observation.

Section 6 contains an immediate application of these ideas, by describing the
rather effortless development of an arithmetically complete axiomatization of a new logic,
which borrows the ∩α operator of [22]. In Section 7 we give a precise definition, based
on the concept of a universe, of Cook's [5] notion of relative completeness, and then
discuss some of the literature regarding this concept, and comment as to the relationship
between it and arithmetical completeness.
2. The Definition of DL and $DL^*$. 

We define regular first-order dynamic logic (DL) as follows: We are given sets of function symbols and predicate symbols, each symbol with a fixed nonnegative arity. We assume the inclusion of the special binary predicate symbol "=" (equality) in the latter set. We denote predicate symbols by $p$, $q$, ..., and $k$-ary function symbols for $k>0$ by $f$, $g$, .... Zeroary function symbols are denoted by $z, x, y, ...$ and are called variables. A term is some $k$-ary function symbol followed by a $k$-tuple of terms, where we restrict ourselves to terms resulting from applying this formation rule finitely many times only. For a variable $x$ we abbreviate $x()$ to $x$, thus $f(g(x),y)$ is a term provided $f$ and $g$ are binary and unary, respectively. An atomic formula is a $k$-ary predicate symbol followed by a $k$-tuple of terms.

We define by simultaneous induction the set $RG$ of first-order regular programs and the set of $DL$-wffs:

1. For any variable $x$ and term $e$, $x=e$ is in $RG$,
2. For any program-free (see below) $DL$-wff $P$, $P?$ is in $RG$,
3. For any $\alpha$ and $\beta$ in $RG$, $(\alpha ; \beta)$, $(\alpha \cup \beta)$ and $\alpha^*$ are in $RG$,
4. Any atomic formula is a $DL$-wff,
5. For any $DL$-wffs $P$ and $Q$, $\alpha$ in $RG$ and variable $x$, $\neg P$, $(P \lor Q)$, $\exists xP$ and $<\alpha>P$ are $DL$-wffs.

A $DL$-wff which contains no occurrence of a program of $RG$ is called program free, or simply a first order formula. Programs of the form indicated in (1) and (2) are called, respectively, (simple) assignments and (simple) tests. We use $\land$, $\Rightarrow$ and $\equiv$ for abbreviations in the standard way and, in addition, abbreviate $\neg \exists x \neg P$ to $\forall x P$, and $<\alpha>\neg P$ to $[\alpha]P$.

The semantics of DL is based on the concept of a state. A state $\mathcal{J}$ consists of a non-empty domain $D$ and a mapping from the sets of function and predicate symbols to the sets of functions and predicates over $D$, such that to a $k$-ary function symbol $f$ (resp. predicate symbol $p$) there corresponds a total $k$-ary function (resp. predicate) over $D$, denoted by $f_J$ (resp. $p_J$). In particular, to a variable there corresponds an element of the domain, and to a 0-ary predicate symbol (propositional letter) a truth value (true or false). The standard equality predicate over $D$ is that which corresponds to the equality symbol (\(=\)). We will sometimes refer to the domain of $\mathcal{J}$ as $D_J$. Observe that the way states are defined no distinction is made between what are normally called variables and constants. These will, however, be defined below for simple universes.

We denote by $\Gamma$ the collection of all possible states, which we call the grand universe. Our semantics will assign to a program $\alpha$ a binary relation $m(\alpha)$ over $\Gamma$, and
to a formula $P$, a subset of $\Gamma$ consisting of those states which satisfy $P$. In the sequel, however, we will be interested in special subsets of $\Gamma$, namely, universes:

A pseudo-universe $U$ is a set of states, all of which have a common domain $D$. A function symbol $f$ (resp. predicate symbol $p$) is called uninterpreted in $U$ if for every state $J \in U$ and for every function $F$ (resp. predicate $P$) over $D$ there exists $J' \in U$ such that $J$ and $J'$ differ at most in the value of $f$ (resp. $p$), which in $J'$ is $F$ (resp. $P$).

Notation: for any function $G: A \to B$, arbitrary element $e$, and $a \in A$, we define $Le/\{e\}G$ to be the function with domain $A$ and range $Bu\{e\}$ giving the same values at points in $A-\{a\}$ as $G$, and such that $G(a) = e$. Thus, the situation described above for uninterpreted $f$ is simply $J=JF/\{J\}J$.

A symbol is called fixed in $U$ if its value is the same in all states of $U$. Thus, "$=$" is fixed in any universe. A universe is a pseudo-universe in which every predicate symbol is fixed, and in which every function symbol is either fixed or uninterpreted. A universe is called simple if the only uninterpreted symbols in it are a designated set of variables. In a simple universe the fixed variables can be called constants following ordinary usage.

The value of a term $e = f(e_1,...,e_k)$ in a state $J$ is defined inductively by

$$e_J = f_J(e_1J,...,e_kJ).$$

We now define by simultaneous induction the binary relation over $\Gamma$ corresponding to a program $\alpha$ of $RGC$, and those states $J$ in $\Gamma$ which satisfy a DL-wff $P$. The relation will be denoted by $m(\alpha)$ and for the latter we write $J \equiv P$. $(J,J)$ being an element of $m(\alpha)$ can be thought of as representing the fact that there exists a computation sequence (or path) of $\alpha$ starting in state $J$ and terminating in $J$. Thus, $J \equiv \langle \alpha \rangle P$ will be seen to be making an assertion about all terminating computations of $\alpha$ starting in state $J$; namely the assertion that the final states of these computations satisfy $P$. Similarly, $J \equiv \langle \alpha \rangle P$ asserts the existence of a terminating computation of $\alpha$ starting in state $J$, which ends in a state satisfying $P$.

1. For any variable $x$ and term $e$,
$$m(x=e) = \{(J,J) | J \vdash e_J = xJ\},$$

2. For any program-free DL-wff $P$,
$$m(P?) = \{(J,J) | J \vdash P\},$$

3. For any $\alpha$ and $\beta$ in $RGC$,
$$m(\alpha;\beta) = m(\alpha) \circ m(\beta), \quad \text{(composition of relations)}$$
$$m(\alpha U \beta) = m(\alpha) U m(\beta), \quad \text{(union of relations)}$$
m(α*) = (m(α))*, \quad \text{(reflexive transitive closure of a relation),}

(4') For an atomic formula p(e1, ..., ek),
J ⊨ p(e1, ..., ek) whenever p_f(e1, ..., ek) is true,

(5') For any DL-wffs P and Q, α in RG and variable x,
J ⊨ \neg P \iff \text{it is not the case that } J \models P,
J ⊨ (P ∨ Q) \iff \text{either } J \models P \text{ or } J \models Q,
J ⊨ \exists x P \iff \text{there exists an element } d \text{ in } D_f \text{ such that } [d / x] J \models P,
J ⊨ <α> P \iff \text{there exists a state } J \text{ such that } (J, J) \in m(α) \text{ and } J \models P.

Note that the only kinds of formulae, whose truth in state J depends possibly upon states other than J, are those containing subformulae of the form ∃xP and <α>P.

In this paper we will primarily be interested in investigating the truth of formulae in a given simple universe U. However, one can see that for some J ∈ U and some assignment x = e, the unique state J such that (J, J) ∈ m(x = e), i.e. the state [e] / x] J, might not be in U at all. We outlaw this phenomenon by adopting, from now on, the convention that in the context of a given universe, the only programs we consider are those in which the variables assigned to (e.g. x in x = e) and the quantified variables (e.g. x in ∃xP) are uninterpreted. Thus, for J ∈ U and for any DL-wff P, the truth of J in P can be seen to depend only on states in U.

We abbreviate (J, J) ∈ m(α) to JαJ, and also take the liberty of writing

J ⊨ <α> P \iff \exists J (JαJ ∧ J ⊨ P), \quad \text{and thus we have also}
J ⊨ [α] P \iff \forall J (JαJ → J ⊨ P).

Given a universe U, we say that a DL-wff P is U-valid (and write P ∈ U) if for every J ∈ U we have J ⊨ P. We say P is valid (⊨P) if it is U-valid for every universe U in which, in line with the above convention, the assigned and quantified variables of P are uninterpreted.

The following are examples of valid DL-wffs:

\[
\begin{align*}
[(x = z ∧ y = u) ; (x = f(x) ∨ y = f(y))] & (x = z ∨ y = u), \\
x = y & \models [(x = f(f(x))) ∨ (y = f(y))] \ (x = y), \\
x = y \models [(x = f(x)) ∨ (y = f(y))] & \ (x = y ∨ (y = f(y))) \ (y = f(y) \ (y = f(y))).
\end{align*}
\]

The first asserts that at most one of the components of u is executed. The second states that the process of repeatedly applying a function composed with itself is a special case of that of repeatedly applying it. The third asserts essentially that the process of achieving a property of x by repeatedly applying f can be simulated in y.
Denote by \( N \) the simple universe of pure arithmetic; i.e., the domain \( D \) is the set of natural numbers, and \( +, \cdot \) and 0 are fixed with their standard interpretations. We freely use standard arithmetical abbreviations such as \( \geq, \gcd, \) etc. The following are \( N \)-valid DL-wffs:

\[
\langle (x+x-1) \rangle x=0, \\
y>0 \lor \langle y=0? \rangle \text{true} \quad \text{(We abbreviate } x=x \text{ to true and } -\text{true to false)}, \\
\langle (x=x' \land y=y' \land x'y>0) \rangle \langle (x\neq y?; (x>y?; x+x-y \lor x<y?; y<y-x)) \rangle x=y? x=x' \gcd(x', y').
\]

The last example asserts that the program inside the diamond, under the assumption that its two inputs are positive integers, terminates and computes the gcd of these inputs. This program can be written in more popular terms as:

\[
\text{while } x\neq y \text{ do if } x>y \text{ then } x=x-y \text{ else } y=y-x \text{ end.}
\]

We adopt the standard definition of a free occurrence of a variable \( x \) in a first order formula \( Q \) to be an occurrence of \( x \) which is not in any subformulae of the form \( \exists x \mathcal{P} \). Any other occurrence is called bound. Also, we define \( Q^e_x \) for variable \( x \) and term \( e \) to be the formula which is obtained from \( Q \) by uniformly renaming all bound variables of \( Q \) which appear in \( e \) and replacing all free occurrences of \( x \) by \( e \).

We now describe our extension of DL, namely DL\(^+\), for dealing with infinite computations. The definitions of DL\(^+\) given in [14] and [9] amount to exactly the same, but are presented differently; in [14] we use a "divergence state", and in [9] computation trees. However, once the present definition or any one of the others has been given, the concepts are well defined and the rest of the treatment is identical.

For any program \( \alpha \in \mathcal{R} \) and state \( I \in \Gamma \), a Boolean constant \( \text{loop}_\alpha \) is added to the vocabulary of DL, giving DL\(^+\). The semantics of \( \text{loop}_\alpha \) are given by induction on the structure of \( \alpha \) as follows:

\[
\text{loop}_{x=e} = \text{false}, \\
\text{loop}_{?} = \text{false}, \\
\text{loop}_{\alpha \lor \beta} = (\text{loop}_\alpha \lor \text{loop}_\beta), \\
\text{loop}_{\alpha; \beta} = (\text{loop}_\alpha \lor \langle x \rangle \text{loop}_\beta).
\]

For \( \alpha^* \) we define \( I \models \text{loop}_\alpha^* \) iff either \( I \models \text{loop}_\alpha \) or there exist states \( I_0, I_1, I_2, \ldots \in I \) such that \( I_0 = I \) and \( (\forall i \geq 0)(I_i \models \text{loop}_\alpha) \).

(We remark that our original notation in [14] had \( dw(\alpha) \) standing for \( \neg \text{loop}_\alpha \). We now prefer the notation \( \text{loop}_\alpha \) which is what is used in [19] and [9].)
Now abbreviate the $\text{DL}^+\text{-wff } (\alpha \vee \text{loop}_\alpha)$ to $\alpha^+\text{P}$, and $\neg\alpha^+\text{P}$ to $[\alpha]^+\text{P}$. Thus we have

$\alpha^+\text{false} \equiv \text{loop}_\alpha$

and

$[\alpha]^+\text{true} \equiv \neg\text{loop}_\alpha$

The intuition is that $\alpha^+\text{false}$ (i.e. $\text{loop}_\alpha$) is true in a state $J$ iff $\alpha$ "can enter an infinite loop"; i.e. "$\alpha$ has a divergence". $[\alpha]^+\text{true}$ is true in $J$ iff $\alpha$ is "divergence-free".

We refer the reader to [211], [12] and [9] for many more details and results concerning $\text{DL}$ and $\text{DL}^+$.

3. Arithmetical Axiomatization of $\text{DL}$ and $\text{DL}^+$.

We would like to supply a syntactic characterization of the $U$-valid $\text{DL}$-wffs and $\text{DL}^+$-wffs for specific universes $U$; namely those which "contain" the simple universe of arithmetical, $\mathcal{N}$. This characterization will be in the form of sound axiom systems for $\text{DL}$ and $\text{DL}^+$ which make explicit use of variables that range over the natural numbers. For any such "arithmetical" universe $A$, we take all $A$-valid first order formulae as further axioms, and show that then the axiom systems are $A$-complete, i.e. that a proof in them does indeed exist for any $A$-valid $\text{DL}$-wff. This property we will term arithmetical completeness.

An arithmetical universe $A$ is a universe in which the domain includes the set of natural numbers, the binary function symbols $+$ and $\cdot$ are fixed and given their standard meanings (addition and multiplication respectively) when applied to the natural numbers in the domain, and 0 is a fixed zeroary-order function symbol interpreted as the natural number "zero". Furthermore there is a fixed unary predicate symbol $\text{nat}$ with the interpretation "$\text{nat}_J(d)$ is true iff $d$ is a natural number", that is, for every state $J$, 

$\{d \in D_J \mid \text{nat}_J(d)\}$

is the set of natural numbers. Thus, we are able to distinguish the natural numbers in the domain from the other elements, and we do not care, say, what the value of $x \cdot y$ is in state $J$ when it is not the case that $\text{nat}_J(x^y)$ holds. Note that one particular arithmetical universe is the simple universe $\mathcal{N}$ of "pure arithmetic" in which $\text{nat}$ is identically true.

Throughout the rest of the paper, $A$ stands for any arithmetical universe, and $L$ for the set of first-order formulae. When talking about arithmetical universes we will often want to use $n, m, \ldots$ to stand for variables ranging only over the natural numbers. We do this by adopting the following convention: Any $L$-wff we will use, in which we have explicitly mentioned, say, the variable $n$ as a free variable, is assumed to be preceded by $\text{nat}(n) \supset$. Thus, for example, $J \vDash (P(n) \supset Q)$ stands for $J \vDash (\text{nat}(n) \supset (P(n) \supset Q))$, asserting
that in state \( I \), \((P(n) \supset Q)\) is true if \( n \) happens to be a natural number. Furthermore, by convention, \( \forall n P(n) \) stands for \( \forall n(n \in \text{nat}(n) \supset P(n)) \), and hence \( \exists n P(n) \) abbreviates \( \exists n(n \in \text{nat}(n) \land P(n)) \).

For any \( \alpha \in \text{RG} \), we let \( \text{var}(\alpha) \) stand for the set of variables appearing in \( \alpha \).

Consider the following axiom system \( P \) for DL:

**Axioms:**

- (A) All tautologies of propositional calculus.
- (B) All \( A \)-valid \( L \)-wffs.
- (C) \( [x \leftarrow e]P \equiv P^x \), for an \( L \)-wff \( P \).
- (D) \( [Q \supset P]P \equiv (Q \supset P) \).
- (E) \( [\alpha ; \beta]P \equiv [\alpha]\lsem \beta \rsem P \).
- (F) \( [\alpha u \beta]P \equiv ([\alpha]P \land [\beta]P) \).

**Inference rules:**

- (G) **Modus Ponens**
  
  \[ \frac{P, P \supset Q}{Q} \]

- (H) \( \frac{P \supset Q}{[\alpha]P \supset [\alpha]Q} \)

- (I) **Invariance**
  
  \[ \frac{P \supset [\alpha]P}{P \supset [\alpha^*]P} \]

- (J) **Convergence**
  
  \[ \frac{P(n+1) \supset \langle \alpha \rangle P(n)}{P(n) \supset \langle \alpha^* \rangle P(0)} \] for an \( L \)-wff \( P \) with free \( n \), \( s.t. \ n \notin \text{var}(\alpha) \).

A \( DL \)-wff \( P \) is said to be **provable** in \( P \), written \( \vdash_P P \), if there exists a finite sequence \( S \) of \( DL \)-wffs, the last one being \( P \), and such that each formula in \( S \) is an axiom (or instance of an axiom scheme), or is obtained from previous formulae of \( S \) by one of the rules of inference.

We now state four theorems which essentially appear in [12], and in more detail in [9].

**Theorem 1** (\( A \)-expressiveness for DL): For any \( DL \)-wff \( P \) there exists an \( L \)-wff \( F_p \) such that \( F_A(P \equiv F_p) \).
Theorem 2 (A-soundness for DL): For any DL-wff P, if $\vdash_P P$, then $\vdash_A P$.

Theorem 3 (Box-completeness): For every $\alpha \in \text{RG}$ and L-wffs R and Q,

if $\vdash_A (R \rightarrow [\alpha]Q)$ then $\vdash_P (R \rightarrow [\alpha]Q)$.

Theorem 4 (Diamond-completeness): For every $\alpha \in \text{RG}$ and L-wffs R and Q,

if $\vdash_A (R \rightarrow <\alpha>Q)$ then $\vdash_P (R \rightarrow <\alpha>Q)$.

Theorem 5 (A-completeness for DL): For any DL-wff P, if $\vdash_A P$ then $\vdash_P P$.

Thus, for any arithmetical universe A, P characterizes the set of A-valid DL-wffs.

Turning to DL+, we remark that one way of supplying an axiomatization of DL+ can be derived from a recent theorem of Winklmann [24], who shows (contrary to our conjecture in [14]) that DL and DL+ are equal in expressive power. His proof provides an algorithm for obtaining, for any $\alpha$, a DL-wff $Q_\alpha$ such that $\vdash (Q_\alpha \equiv <\alpha>^{+}\text{false})$. Thus, one could add a version of this algorithm to $P$ as a rule for transforming $<\alpha>^{+}\text{false}$ into $Q_\alpha$, and then the additional axiom,

$$<\alpha>^{+}P \equiv (<\alpha>P \lor <\alpha>^{+}\text{false})$$

would render this an arithmetically complete axiomatization of DL+. However, $Q_\alpha$ is a rather complicated formula and was designed in order to prove a "power of expression" result. The goals of axiomatization, and arithmetical axiomatization in particular, are different. Here we are interested in concise, elegant and well-structured axiom systems for the purpose of proof. Consequently, we will concentrate on our original axiomatization, given in [14].

Augment P with the following, to obtain $P^+$:

Axioms:

(K) $[x\in E]^+\text{true},$
(L) $[Q]^+\text{true},$
(M) $[\alpha ; \beta]^+\text{true} \equiv [\alpha]^+[\beta]^+\text{true},$
(N) $[\alpha \lor \beta]^+\text{true} \equiv ([\alpha]^+\text{true} \land [\beta]^+\text{true}),$
(O) $[\alpha]^+P \equiv ([\alpha]P \land [\alpha]^+\text{true}),$

Inference rules:

(P) Finiteness $P(n+1) \Rightarrow [\alpha]^+P(n), \neg P(0)$

\[ P(n) \Rightarrow [\alpha]^+\text{true} \]

for an L-wff P with free n, s.t. n \notin \text{var}(\alpha),
(Q) Divergence  \[ P \Rightarrow \alpha \Rightarrow P \]
\[ P \Rightarrow \alpha \Rightarrow \text{false} \]

Here too we have:

**Theorem 6** (A-expressiveness for DL⁺): For any DL⁺-wff P there exists an L-wff \( F_\alpha \) such that \( F_A (P \equiv F_\alpha) \).

**Theorem 7** (A-soundness for DL⁺): For any DL⁺-wff P, if \( F_P \vdash P \), then \( F_A \vdash P \).

**Theorem 8** (Box⁺-completeness): For every \( \alpha \in \mathcal{RG} \) and L-wffs R and Q,
\[ \text{if } F_A (R \Rightarrow \alpha \Rightarrow Q) \text{ then } F_P \vdash (R \Rightarrow \alpha \Rightarrow Q). \]

**Theorem 9** (Diamond⁺-completeness): For every \( \alpha \in \mathcal{RG} \) and L-wffs R and Q,
\[ \text{if } F_A (R \Rightarrow \alpha \Rightarrow Q) \text{ then } F_P \vdash (R \Rightarrow \alpha \Rightarrow Q). \]

**Theorem 10** (A-completeness for DL⁺): For any DL⁺-wff P, if \( F_A \vdash P \), then \( F_P \vdash P \).

We now set ourselves out to find the thread connecting the previous results.

4. **The Theorem of Completeness.**

In this section we prove a general theorem which can be seen to provide the inductive step used in the proofs of all the specific completeness results of [12], [14] and [39], and in particular, of the present Theorems 5 and 10. Also, the proofs of arithmetical completeness results for any DL-like logics which fall under the general criterions of this theorem can be considerably shortened by appealing to it. Effort can then be devoted entirely to establishing the "base" of this induction, as illustrated in the next section.

As above, we denote the set of first-order formulae by L. Assume we are given a universe U, a set K, and a functional

\[ M : K \times 2^U \rightarrow 2^U. \]

The M-extension of L, \( L(M) \), is defined to be the following language which is L augmented with one formation-rule:

1. Any atomic formula is in \( L(M) \),

\[ \]
(2) For any \(k \in K\), variable \(x\) and \(L(M)\)-wffs \(P\) and \(Q\), 
\(-P\), \((P \lor Q)\), \(\exists xP\) and \((M_k)P\) are \(L(M)\)-wffs.

The semantics of \(L(M)\) are defined with \(\mathcal{S} = (M_k)P\), holding whenever \(\mathcal{S} \in M(k, \{\phi | \mathcal{S} = \phi\})\), and with the other clauses receiving their standard meanings.

Some intuition might be gained at this point by noticing that if \(K\) is taken to be the class \(RG\) of regular programs over assignments and tests, and \((M_\alpha)P\) is interpreted as \(\langle \alpha \rangle P\), then \(L(M)\) is in fact precisely \(DL\).

We now define some important concepts to be used in the sequel:

We say that \(L\) is \(U\)-expressive for \(L(M)\) if for every \(L(M)\)-wff \(P\) there exists an \(L\)-wff \(Q\) such that \(\models_U (P \equiv Q)\).

An axiom system \(P(M)\) for \(L(M)\) is any set of axioms (or axiom schemas) and inference rules over \(L(M)\). Provability of an \(L(M)\)-wff \(P\) in \(P(M)\) is defined in the standard way and is denoted by \(\models_{P(M)} P\). \(P(M)\) is said to be \(U\)-sound if all the axioms are \(U\)-valid and all the rules of inference preserve \(U\)-validity. Note then, that whenever \(\models_{P(M)} R\), we have \(\models_U R\).

\(P(M)\) is said to be \(propositionally complete\) if all instances of tautologies of propositional calculus are theorems of \(P(M)\) and \(modus ponens\) is in the set of inference rules.

It is said to be \(U\)-complete if for every \(L(M)\)-wff \(R\), whenever \(\models_U R\), we have \(\models_{P(M)} R\).

**Theorem 11 (Theorem of Completeness):** For any universe \(U\), \(M\)-extension \(L(M)\) of \(L\), a \(U\)-sound axiom system \(P(M)\) for \(L(M)\) is \(U\)-complete whenever the following hold:

1. \(P(M)\) is propositionally complete,
2. \(L\) is \(U\)-expressive for \(L(M)\),
3. For any \(k \in K\) and \(L(M)\)-wffs \(R\) and \(Q\),
   \[
   \text{if } \models_{P(M)} (R \Rightarrow Q) \text{ then } \models_{P(M)} ((M_k)R \Rightarrow (M_k)Q),
   \]
4. For any \(k \in K\) and \(L\)-wffs \(R\) and \(Q\),
   a) if \(\models_U R\) then \(\models_{P(M)} R\),
   b) if \(\models_U (R \Rightarrow (M_k)Q)\) then \(\models_{P(M)} (R \Rightarrow (M_k)Q)\), and
   c) if \(\models_U (R \Rightarrow \neg (M_k)Q)\) then \(\models_{P(M)} (R \Rightarrow \neg (M_k)Q)\).

**Proof:** We have to prove that if \(P\) is an \(L(M)\)-wff such that \(\models_U P\), then \(\models_{P(M)} P\).

By the propositional completeness of \(P(M)\) we can assume that \(P\) is given in conjunctive normal form, and we proceed by induction on the sum of the number of appearances of \(M\) and the number of quantifiers in \(P\). Assume the theorem holds for any formula with \(n-1\) or less appearances of \(M\) and quantifiers. If \(P\) is of the form \(P_1 \land P_2\), then we have \(\models_U P_1\) and \(\models_U P_2\) and...
P2, so that we can restrict our attention to a single disjunction. Without loss of
generality we can, therefore, assume that P is of one of the forms:

\[ P_{1\lor}(M_k)_{P2}, \quad P_{1\lor}-(M_k)_{P2}, \quad P_{1\lor}3x_{P2} \quad \text{or} \quad P_{1\lor}3x_{P2}, \]

where \( k \in K \) and \( P_1 \) and \( P_2 \) each have \( n-1 \) or less appearances of \( M \) and quantifiers. Let us
use \( \rho \) to denote \( (M_k), -(M_k), 3x \) or \( \neg 3x \) according to which is the case.

\( L \) is expressive for \( L(M) \), and so for any \( L(M) \)–wff \( Q \) there is some \( L \)–wff \( F_Q \) which
is equivalent to \( Q \). We have then \( \models_U (\neg F_{P1} \Rightarrow \rho F_{P2}) \). Now, using assumption (4)
(since \( F_{P1} \) and \( F_{P2} \) are \( L \)–wffs), we also have

\( (+) \quad \models_P(M) (\neg F_{P1} \Rightarrow \rho F_{P2}). \)

Now surely, by the definition of \( F_{P1} \) and \( F_{P2} \), we have \( \models_U (\neg P_1 \Rightarrow \neg F_{P1}) \) and \( \models_U 
(\neg F_{P2} \Rightarrow P_2) \). Both these last formulae have less than \( n \) appearances of \( M \) and
quantifiers, and hence by the inductive hypothesis

\( (++) \quad \models_P(M) (\neg P_1 \Rightarrow \neg F_{P1}) \quad \text{and} \quad
\models_P(M) (F_{P2} \Rightarrow P_2). \)

By assumption (3) or (4a) (depending on whether \( \rho \) is an appearance of \( M \) or a quantifier)
together with the propositional completeness, we obtain from the latter

\( (+++) \quad \models_P(M) (\neg F_{P2} \Rightarrow \rho P_2). \)

From \((+)\), \((++)\) and \((+++)\) we get, using propositional reasoning, \( \models_P(M) (\neg P_1 \Rightarrow \rho P_2) \),
or \( \models_P(M) (P_1 \lor \rho P_2). \)

Now, take \( A \) to be any arithmetical universe and \( K \) to be the set \( \mathcal{R} \mathcal{C} \) of regular
programs over assignments and tests. Take \( \mathcal{M} \) to be the "diamond" operator, or more
precisely define \( \mathcal{F}(M_\alpha)_{P} \) iff \( \mathcal{F} < \alpha \alpha P \). Certainly then, \( L(M) \) is simply \( DL \). Furthermore,
note that in this framework taking \( P(M) \) to be the system \( P \), we see that Theorem 2
establishes the \( \mathcal{U} \)–soundness required in Theorem 11, axiom scheme (A) and rule (G) satisfy
requirement (1) of the theorem, Theorem 1 satisfies requirement (2), rule (H) (or, more
precisely, an easily derived version of it) satisfies requirement (3), and Theorems 3 and
4, together with axiom scheme (B), satisfy requirement (4). Consequently, Theorem 5 can
be seen to be a corollary of Theorem 11, using Theorems 1–4.

Turning to \( DL^* \), it is quite easy to see that Theorem 11 can be extended to deal
with "double extensions" of $L$; i.e. for the case where there are two functionals $M$ and $M'$. We omit the precise rephrasing of the theorem for this case, but note that taking $(M^*\alpha)P$ to be $<\alpha>P$ as above and $(M^*\alpha)P$ to be $<\alpha>^+P$, the $(M_M^*)$-extension, $L(M,M^*)$, is simply $DL^+$. Here too, Theorem 10 can be seen to be a corollary of Theorem 11, using Theorems 6–9.

Theorem 11 captures the inductive step used in proving all the completeness results of [12], [14] and [9]. These include results similar to Theorems 5 and 10 for the extensions of $DL$ and $DL^+$ for dealing with recursive programs, CFDL and CFDL+$^*$. In these cases the set $K$ is taken to be the set $CF$ of context-free programs over assignments and tests.

Note now that requirements (1), (3) and (4a) of Theorem 11 were satisfied in $P$ and $P^+$ by simply including, in a straightforward manner, axiom schemes (A), and (B) and rules (C) and (H). In the next section, however, we try to see how $P$ and $P^+$ were made to satisfy requirements (4b) and (4c).

5. One-program Completeness Results.

In this section we describe how $P$ and $P^+$ were designed so that the basic completeness results, i.e., Theorems 3,4,8 and 9, could be established in [12] and [14].

Let us first consider $DL$. Theorem 3 can be seen to be essentially Cook's [5] theorem concerning Hoare's [15] axiom system for proving the partial correctness of regular deterministic programs (see also [11], and Section 3.3 of [9]). Hoare's approach was to construct the system in such a way that the program $\alpha$ in the formula $R[\alpha]Q$ (which he wrote as $R\{\alpha\}Q$) is "decomposed" one step at a time, based on the structure of $\alpha$. The idea was to supply one rule of inference for every formation rule of the programming language, enabling one to conclude a property of a program by establishing similar properties of its immediate components. Completeness is then proved by induction on the structure of $\alpha$, showing that at each step the appropriate rule could indeed be applied.

This appealing approach is imitated in our systems $P$ and $P^+$. One can show that axioms (C)–(F), with the addition of (K)–(O) for $DL^+$ together with the ability (made possible by propositional completeness) to reason about $<\alpha>$ and $<\alpha>^+$ using $[\alpha]$ and $[\alpha]^+$, serve to enable the carrying out of this decomposition for all the program constructs in RC but one: $\alpha^*$.

Finding the appropriate axioms or rules for proving the properties of $\alpha^*$ is what we might now call the heart of the problem. Let us assume that we can find a sound rule of inference having (a) $R[\alpha^*]Q$ as its conclusion, (b) programs of complexity "at most" $\alpha$ in its premises, and (c) the property that whenever $R[\alpha^*]Q$ is $A$–valid, then one can
establish the $A$-validity of its premises. Then the task of proving the $A$-validity of $R \Rightarrow [\alpha^*]Q$ can always be reduced to that of proving the "simpler" premises. This fact would then serve as the $\alpha^*$ part of the proof of Theorem 3 "in the spirit" of Hoare [15].

Similarly, such a rule for $R \Rightarrow <\alpha^*>Q$ would establish Theorem 4, and these two would, as shown above, complete the proof of the general completeness result, Theorem 5. In the same way, having axiom (O) at hand, rules for $R \Rightarrow [\alpha^*]^+true$ and $R \Rightarrow <\alpha^*>^+false$ can be seen to be sufficient for establishing Theorem 10.

We will show that these four tasks, resulting in rules (I), (J), (P) and (Q), are quite easy and are dual to one another in a rather interesting way.

We will work on $[\alpha^*]$ and $<\alpha^*>$ simultaneously, and later exhibit the analogous process for $[\alpha^*]^+$ and $<\alpha^*>^+$.

Let us take a look at the concept involved:

$[\alpha^*]Q$  \hspace{5cm} $<\alpha^*>Q$.

We have the "arithmetical equivalent" of our concept:

$\forall n [\alpha^n]Q$  \hspace{5cm} $\exists n <\alpha^n>Q$.

Note that for a given $n$ and state $J \leq A$, we can prove

$J \models [\alpha^n]Q$  \hspace{5cm} $J \models <\alpha^n>Q$,

by finding some formula $P$ with a free variable $n$, $n \notin \text{var}(\alpha)$, such that the following hold:

$J \models P(n)$  \hspace{5cm} $J \models <\alpha^n>Q$;

$J \models \forall m (P(m+1) \Rightarrow [\alpha]P(m))$  \hspace{5cm} $J \models \forall m (P(m+1) \Rightarrow <\alpha>P(m))$;

$J \models (P(0) \Rightarrow Q)$  \hspace{5cm} $J \models (P(0) \Rightarrow Q)$.

The reason is that the above establishes, by induction on $n$, that

$J \models [\alpha][\alpha]...[\alpha]Q$  \hspace{5cm} $J \models <\alpha><\alpha>...<\alpha>Q$,

with $n$ occurrences of $\alpha$, which, by

$J \models ([\alpha][\beta]Q \equiv [\alpha ; \beta]Q)$  \hspace{5cm} $J \models (<\alpha><\beta>Q \equiv <\alpha ; \beta>Q)$,

gives

$J \models [\alpha^n]Q$  \hspace{5cm} $J \models <\alpha^n>Q$.

Thus, if we can prove

$J \models (R \Rightarrow \forall n P(n))$  \hspace{5cm} $J \models (R \Rightarrow \exists n P(n))$;

instead of $J \models P(n)$, we will have established

$J \models (R \Rightarrow \forall n [\alpha^n]Q)$  \hspace{5cm} $J \models (R \Rightarrow \exists n <\alpha^n>Q)$.

And so, being able to prove

$J \models (R \Rightarrow \forall n P(n))$  \hspace{5cm} $J \models (R \Rightarrow \exists n P(n))$;
\( \vdash_A (P(m+1) \implies [\alpha]P(m)) \)
\( \vdash_A (P(0) \implies Q) \)
will result in the establishment of
\( \vdash_A (R \implies [\alpha^*]Q) \)
\( \vdash_A (P(m+1) \implies <\alpha> P(m)) \)
\( \vdash_A (P(0) \implies Q) \)
\( \vdash_A (R \implies <\alpha^*> Q) \).

This gives rise to the A-sound rule of inference
\[
\begin{array}{c}
\text{R} \implies \forall n P, \ P' \implies [\alpha] P, \ P^0 \implies Q \\
\hline
\text{(*)}
\end{array}
\]
\( \vdash_A [\alpha^*] Q \)
\( \vdash_A <\alpha^*> Q \)

where \( P' \) abbreviates \( P(n+1) \), and \( P^0 \) abbreviates \( P(0) \).

Now, in order for our required completeness property to hold, we must show that whenever
\( \vdash_A [\alpha^*] Q \)
\( \vdash_A <\alpha^*> Q \)
is A-valid, there exists a \( P(n) \) satisfying the premises of the rule. This, however, is true: take \( P(n) \) to be simply some first-order formula which for any \( n \) is \( A \)-equivalent to
\( [\alpha^n] Q \)
\( <\alpha^n> Q \)

itself. Such a formula exists as shown in [9]. In other words we have
\( \vdash_A \forall n (P(n) \equiv [\alpha^n] Q) \)
\( \vdash_A \forall n (P(n) \equiv <\alpha^n> Q) \).

Thus in fact, we have found sound and complete rules for \( [\alpha^*] \) in DL, and we could have stopped here. However, we would like to go a step further. Using the observation that for the \( [\alpha^*] \) case we have the duality principle \((R \implies [\beta] Q) \equiv (<\beta^-> R \implies Q)\), where \( <\alpha^-> \) is the inverse of \( \alpha \), defined such that \( m(\alpha^-) = \{(I,J) | J \alpha J \} \) (see [21]), we note that we can then write down an "ascending" induction rule, which is constructed symmetrically to the "descending rule" above:

\( \vdash_A \forall n (P(n) \equiv [\alpha^-]^n Q) \)
\( \vdash_A \forall n (P(n) \equiv <\alpha^->^n Q) \).

which by the duality principle is really
\( \vdash_A \forall n (P(n) \equiv [\alpha] Q) \)
\( \vdash_A \forall n (P(n) \equiv <\alpha^*> Q) \)

and can be seen to be complete by taking \( P(n) \) to be \( <(\alpha^-)^n> R \). (The A-expressiveness of \( L \) is retained in the presence of the inverse operator \( ^- \).) Now, because of the universal
character of \([\alpha^*]\) and the fact that \(\forall n[\alpha^n]P \supset [\alpha]\forall n[\alpha^n]P\), or \([\alpha^*]P \supset [\alpha][\alpha^*]P\), we are able to collapse the descending and ascending rules \((*)\) and \((***)\) into one sound and complete rule by eliminating the \(n\)–indexing:

\[
\begin{align*}
R \supset P, & \\
& P \supset [\alpha]P, & P \supset Q \\
\hline
R \supset [\alpha^*]Q
\end{align*}
\]

The premises of this rule are made \(A\)–valid when the conclusion is, by both \(\forall n[\alpha^n]Q\) and \(\exists n(\alpha^n)^n \supset R\), or essentially, by both \([\alpha^*]Q\) and \((\alpha^n)^n \supset R\). (Note that \(*\) replaces \(n\) in these satisfying formulae for \((*)\) and \((***)\).) Hence, we have arrived at the weakest box–antecedent and strongest box–consequent (see [12]), both of which can be used as the loop invariant \(P\), stemming from the descending and ascending induction rules respectively.

Since there is no strongest diamond–consequent (see [12]), it turns out that there is, for the \((\alpha^*)\) case, only the uninteresting ascending rule

\[
\begin{align*}
R \supset P^0, & \\
& [\alpha^*]P \supset P', & \forall nP \supset Q \\
\hline
\forall n[(\alpha^*)^n]R \supset Q
\end{align*}
\]

the consequent of which is not \(A\)–equivalent to \(R \supset (\alpha^*)Q\).

We are left, then, with

\[
\begin{align*}
R \supset P, & \\
& P \supset [\alpha]P, & P \supset Q \\
\hline
R \supset [\alpha^*]Q
\end{align*}
\]

\[
\begin{align*}
R \supset \exists nP, & \\
& P' \supset (\alpha^*)P, & P^0 \supset Q \\
\hline
R \supset (\alpha^*)Q
\end{align*}
\]

The final step is simply pruning the arms and legs of these rules in order to make them more concise, noting, e.g., that from having proved \(R \supset \exists nP\), \(P^0 \supset Q\) and \(\forall n(P \supset (\alpha^*)P^0)\), we can deduce \(R \supset (\alpha^*)Q\) using validities of first order logic (included as axioms in (A)). We arrive, therefore, at our rules of invariance \((I)\) and convergence \((J)\)

\[
\begin{align*}
P \supset [\alpha]P \\
P \supset [\alpha^*]P
\end{align*}
\]

\[
\begin{align*}
P' \supset (\alpha^*)P \\
P \supset (\alpha^*)P^0
\end{align*}
\]

and have in effect "developed" the invariant assertion method for partial correctness which is captured by rule \((I)\) in parallel to the analogous method for its dual.
We would now like to present an identical chain of thought for \( DL^+ \), resulting in a surprisingly similar pair of rules. There we also see two concepts, one of universal and the other of existensial character, with the universal \( \langle \alpha^* \rangle_B \text{false} \) giving rise to an index-free rule with strongest and weakest satisfying predicates. We will do this by summarizing the above discussion and giving the main points for \( \lceil \alpha^* \rceil \) and \( \langle \alpha^* \rangle \) again, together with the analogues for \( \lceil \alpha^* \rceil_B \) and \( \langle \alpha^* \rangle_B \). First, we restate here the results from \[14]:

**Theorem 12.**

1. \( \lceil \alpha^* \rceil_B \text{true} = \exists n \lceil \alpha^n \rceil_B \text{false} \land \lceil \alpha^* \rceil_B \text{true} = \exists n \lceil \alpha^n \rceil_B \text{false} \),
2. \( \langle \alpha^* \rangle_B \text{false} = \forall n \langle \alpha^n \rangle_B \text{true} \lor 
\langle \alpha^* \rangle_B \langle \alpha \rangle_B \text{false} = \forall n \langle \alpha^n \rangle_B \text{true} \).

The intuitive meaning of, say, (2) being that a divergence in \( \alpha^* \) is due either to being able to run \( \alpha \) for ever, or to being able to run \( \alpha \) some number of times and then have \( \alpha \) itself converge.

The concept involved is

<table>
<thead>
<tr>
<th>( \lceil \alpha^* \rceil_B \text{true} )</th>
<th>( \langle \alpha^* \rangle_B \text{false} )</th>
</tr>
</thead>
</table>

The arithmetical equivalent is

<table>
<thead>
<tr>
<th>( \forall n \lceil \alpha^n \rceil_B \text{true} )</th>
<th>( \exists n \langle \alpha^n \rangle_B \text{true} )</th>
</tr>
</thead>
</table>

The descending rule is

<table>
<thead>
<tr>
<th>( R \vdash \forall n P, P \vdash [\alpha] P, P^0 \vdash Q )</th>
<th>( R \vdash \exists n P, P \vdash [\alpha] P, P^0 \vdash Q )</th>
</tr>
</thead>
<tbody>
<tr>
<td>( R \vdash \forall n \lceil \alpha^n \rceil_B \text{true} )</td>
<td>( R \vdash \exists n \langle \alpha^n \rangle_B \text{true} )</td>
</tr>
</tbody>
</table>

<table>
<thead>
<tr>
<th>( R \vdash \exists n P, P \vdash [\alpha] P, P^0 \vdash false )</th>
<th>( R \vdash \forall n P, P \vdash [\alpha] P, P^0 \vdash true )</th>
</tr>
</thead>
<tbody>
<tr>
<td>( R \vdash \exists n \lceil \alpha^n \rceil_B \text{false} )</td>
<td>( R \vdash \forall n \langle \alpha^n \rangle_B \text{true} )</td>
</tr>
</tbody>
</table>
The premises are satisfied (when the consequent is A-valid) by taking \( P \) to be A-equivalent to

\[
\begin{array}{|c|c|}
\hline
[\alpha^n]Q & \langle \alpha^n \rangle Q \\
[\alpha^n]^+false & \langle \alpha^n \rangle ^+true \\
\hline
\end{array}
\]

(We could have stopped here too; the above rules are sound and complete, and will enable the \( DL^+ \)-completeness theorem to go through. We continue however, as we did above.)

The ascending rule is

\[
\begin{array}{|c|}
\hline
R \supset \mathcal{P}^0 , P \supset [\alpha] \mathcal{P}' , \exists \mathcal{P} \supset Q \\
\hline
R \supset [\alpha^+]Q \\
no rule \\
no rule \\
\hline
R \supset [\alpha^+]^+false \\
R \supset \mathcal{P}^0 , P \supset <\alpha^+]^+P' , \exists \mathcal{P} \supset true \\
\hline
\end{array}
\]

The premises of this rule are satisfied by

\[
\begin{array}{|c|}
\hline
\langle (\alpha^-)^n \rangle R \\
no rule \\
no rule \\
\hline
\langle (\alpha^-)^n \rangle R \land [\alpha^+]^+false. \\
\hline
\end{array}
\]

The unified rule is

\[
\begin{array}{|c|}
\hline
R \supset \mathcal{P} , P \supset [\alpha] \mathcal{P} , P \supset Q \\
\hline
R \supset [\alpha^+]Q \\
no rule \\
no rule \\
\hline
R \supset [\alpha^+]^+false \\
R \supset \mathcal{P} , P \supset <\alpha^+]^+P , P \supset true \\
\hline
\end{array}
\]

The premises of this rule are satisfied by both

\[
\begin{array}{|c|}
\hline
[\alpha^+]Q \text{ and } \langle (\alpha^-)^n \rangle R \\
no rule \\
\hline
no rule \\
\langle \alpha^+]^+false \text{ and } \langle (\alpha^-)^n \rangle R \land [\alpha^+]^+false. \\
\hline
\end{array}
\]
The final "pruned" rule is

<table>
<thead>
<tr>
<th>P⇒[α]P</th>
<th>P′⇒&lt;α&gt;P</th>
</tr>
</thead>
<tbody>
<tr>
<td>P⇒[α*]P</td>
<td>P⇒&lt;α*&gt;P^0</td>
</tr>
<tr>
<td>P⇒[α*]P^0</td>
<td>P⇒&lt;α*&gt;+P</td>
</tr>
<tr>
<td>P⇒[α*]true</td>
<td>P⇒&lt;α*&gt;false</td>
</tr>
</tbody>
</table>

The name given to the construct used in a proof (i.e. the P involved) is

<table>
<thead>
<tr>
<th>invariant</th>
<th>convergent</th>
</tr>
</thead>
<tbody>
<tr>
<td>??</td>
<td>divergent (we suggest)</td>
</tr>
</tbody>
</table>

We would appreciate suggestions on names for the "??".

We would like the reader to consider the virtues of conducting this reasoning for the language of regular expressions over assignments and tests. Consider how much more obscure the observations of this section would have been if we were to reason about, say, the while statement, instead of about α*. In our opinion α* captures the raw essence of iterating in programming languages, just as α∪β captures the essence of branching and α;β the essence of sequencing. For the programming language designer who is interested in a deterministic language or in a more "disciplined" nondeterministic one, we can recommend means of restricting the generality of these constructs (e.g. if-then-else and while, or Dijkstra's [6] IF and DO, etc.). Note how the invariant assertion method of Floyd [7], as described by Hoare's while rule [15] (see also Section 3.3 of [9]), has been shown to fall out of this general pattern of arithmetically complete rules as a special case.

A word about recursion. In [9] and [10], we have developed an arithmetically complete system for an augmented programming language which has μXτ(X) as an additional program construct (see, say, [23]). The basic idea there is to develop an analogy between μXτ(X) and τ^n(false?) on one hand, and α* and α^n on the other. That is, α* being \( \bigcup_{n=0}^{\infty} \alpha^n \), enabled us to "count" how many times we iterated α, and to use this counting in the construction of our rules. Similarly, we are using the fact that (for continuous τ's) μXτ(X) is \( \bigcup_{n=0}^{\infty} \tau^n(false?) \) to "count" how many times we recurred with τ, and we have obtained similar results. Thus, it seems that working with α* is showing its advantages in applying these ideas to a more complicated construct.
6. An Application.

We would now like to illustrate the usefulness of the step-by-step development of the \( \alpha^* \) rules for DL and \( DL^* \) in Section 5 by considering another extension of DL, namely ADL.

We add inductively, for any program \( \alpha \) and ADL-wff \( P \), the formula \((\cap \alpha) P\), defining \( \models (\cap \alpha) P \) iff for all \( n \), \( \models (\alpha^n) P \). Thus, we might write \( \models_A (\cap \alpha) P \equiv \forall n (\alpha^n) P \). \((\cap \alpha) P \) is a construct used in the (deterministic) algorithmic logic of Salwicki [22]. Theorem 12 supplies an easy translation of \( DL^+ \) into ADL. In fact, Winklmann’s [24] theorem shows that ADL, \( DL^+ \) and DL are all equivalent in expressive power. However, as was the case for \( DL^+ \), we are interested in an arithmetically complete axiom system for ADL, and therefore simply follow the lines of the previous section. First we must have

**Theorem 13** (A-expressiveness for ADL): For any ADL-wff \( P \), there exists an L-wff \( F_P \), such that \( \models_A (P = F_P) \).

**Proof.** Easy inductive proof using the definition \((\cap \alpha) Q \equiv \forall n (\alpha^n) Q \).

**Lemma 14.** For any program \( \alpha \) and L-wffs \( R \) and \( Q \),

\[
\text{if } \models_A (R \Rightarrow Q) \text{ then } \models_A ((\cap \alpha) R \Rightarrow (\cap \alpha) Q).
\]

**Proof.** We omit this slightly tedious but nevertheless straightforward proof.

Having Lemma 14 at hand, we add the following rule to \( P \):

\[
\begin{array}{c}
R \Rightarrow Q \\
\hline
(\cap \alpha) P \Rightarrow (\cap \alpha) Q
\end{array}
\]

We also add the rules:

\[
\begin{array}{c}
R \Rightarrow P(n) \\
P(n+1) \Rightarrow (\alpha) P(n) \\
P(0) \Rightarrow Q
\end{array}
\]

For an L-wff \( P \) with free \( n \), s.t. \( n \notin \text{var}(\alpha) \),

\[
\begin{array}{c}
P(n+1) \Rightarrow [\alpha] P(n)
\end{array}
\]

For an L-wff \( P \) s.t. \( n \notin \text{var}(\alpha) \),

\[
\begin{array}{c}
P(n) \Rightarrow (\cap \alpha) \Rightarrow P(0)
\end{array}
\]

\((S)\) and \((T)\) are obtained from the following rules, which in turn follow easily from considerations similar to those described in Section 5:
\[ R \supseteq \forall n P, \ P \supseteq \alpha P, \ P_0 \supseteq Q \]
and
\[ R \supseteq \exists n P, \ P \supseteq [\alpha]P, \ P_0 \supseteq Q \]

\[ \forall n \langle \alpha^n\rangle Q \]
\[ \exists n [\alpha^n] Q \]

We do not know of a duality principle, or of any other way for doing away with the indices in rule (S). Denoting the resulting axiom system by \( P(\cap) \), we have

**Theorem 15 (A–soundness and completeness for ADL):** For any ADL–wff \( P \), \( \models_A P \iff \models P(\cap) P \).

**Proof.** Apply Theorem 11, proving assumptions (4b) and (4c) of that theorem using the above rules, by showing that their premises are satisfied respectively by \( \langle \alpha^n\rangle Q \) and \( [\alpha^n] Q \).

**7. Relative Completeness.**

The approach to axiomatization taken in this paper is closely related to, and in fact is derived from, Cook's [5] notion of relative completeness. In this section we take up the task of comparing the two approaches.

As we indicated in the previous section, Hoare [15] introduced an axiom system for proving the partial correctness of deterministic programs. For the sake of this discussion we can in fact think of the subsystem of \( P \) consisting of axioms (A), (C)–(F) and rules (G) and (I) as Hoare's system, and can denote it by \( \mathcal{H} \) (see Section 3.3 of [39]). Cook [5] investigated the question of completeness of Hoare's system and managed to formalize what seems to be the intuitive way in which people prove correctness (partial in this case) of programs in line with the method suggested by Floyd [7] and Naur [20]. Cook separated the reasoning about the program from the reasoning about the underlying language, making a distinction between proving say, \( \{x+1\} x=1 \), and proving \( (x>0 \Rightarrow x \geq 0) \). The former still requires some program–oriented manipulation in order to turn it into a first–order formula, whereas the second does not. Thus, Cook's idea was to supply Hoare's system with a generous oracle which had the ability to answer questions concerning the truth of first order formulae. In this way he was able to shift concentration to Hoare's rules themselves, which were to serve as a tool for performing a step–by–step transformation of partial correctness assertions (of the form \( P \supseteq [\alpha] Q \)) into equivalent first–order formulae. The truth of the latter is then checked using the oracle.

The way we formalize this approach is, as we have done, to represent the oracle simply as the collection of the true first–order formulae that we are interested in, taken as further axioms. Thus we arrive at the following formal definition using the terminology we have developed: Assume we are given a language \( L' \) which includes all
first-order formulae as wffs; thus L is part of L'. Assume AX is a sound axiom system for L', and denote by AX_U the system AX ∪ \{P | P ∈ L and U_P P\}. In other words, AX_U is AX augmented with all the U-valid first-order formulae as further axioms. AX is said to be complete for L' relative to L if for every universe U such that L is U-expressive for L', AX_U is U-complete for L' (every U-valid L'-wff is provable in AX_U).

**Theorem 16 (Cook):** Denote by L_H the language \{R →_H α | R and Q are L-wffs\}. Then, H is complete for L_H relative to L.

The proof is in fact identical to that of our Box-completeness theorem (Thm. 3).

Now, if we restrict ourselves to languages L' such that for any arithmetical universe A, L is A-expressive for L' (as is the case, for example, when L' is taken to be any one of DL, DL^+, CFDL, CFDL^+ etc.), we note that arithmetical completeness is a special case of relative completeness; we do not require that AX_U be U-complete for all universes U which make L U-expressive for L', but only that that be the case for any arithmetical universe. Consequently then, in AX itself we may use symbols in ways which take their standard interpretation for granted. This is the flavor of the usage of n, * and 0 in the rules (J) and (P), as well as in the rules for (∩α) in Section 6.

Let us look at some related work. Although coming earlier than [5] and mentioning neither expressiveness nor Hoare-like systems, Theorem 1 of [18], Theorem 2(1) of [33], and the part of [1] which considers regular programs, include the observations needed for obtaining relative completeness theorems for Floyd's inductive assertion method, and would carry over to suitable versions of Hoare's system.

Extensions of Cook's result to cover full mutually recursive procedures appear in [8] and in [13], and both utilize a method for "freezing" the input variables upon entry to a procedure. (These systems are subsumed by the axiomatization of context-free DL in [10] and [93].)

This flurry of "positive" research led inevitably to a counter-effort of "negative" research aimed at proving incompleteness results which indicate when Hoare-like systems are doomed to be incomplete even in the relative sense of Cook. The first notable result in this direction is that of Wand [23], who shows essentially that it is not the case that L is U-expressive for every universe U. Thus Wand shows that there exist universes U such that AX_U is not U-complete for L_H. More recently, Lipton [16] claims to have proved the following very interesting characterization of these "good" universes: L is U-expressive for L_H i f f U is an arithmetical universe or a universe with a finite domain (call the latter a finite universe). Thus according to this claim the only universes for which a Hoare-like system can be relatively complete are the arithmetical
ones and the finite ones. So Cook's \([5] \) requirement boils down to requiring that \(\mathcal{AX}_U \) be \(U\)-complete for these two kinds of universes.

The finite universes, however, cause trouble: Clarke \([4] \) has shown that introducing (into the programming language in which the programs of \(\mathcal{L}_H \) are written) various programming concepts such as procedures as parameters or coroutines, in the presence of recursion and other reasonable mechanisms, prevents the possibility of obtaining relatively complete axiom systems. The argument in \([4] \) is based on the fact that the first order language \(\mathcal{L} \) is \(U\)-expressive for \(\mathcal{L}_H \) for any finite universe \(U\). The incompleteness results are then established by showing that these complex programming languages have an undecidable halting problem over finite domains, and hence the set of diverging programs is not r.e., a fact which would contradict the existence of any relatively complete Hoare-like axiom system for such a language (the existence of one implying that, in particular, the set of valid formula of the form \(true \Rightarrow \alpha \) fals is r.e.). Hence, the essence of Clarke's results lies in the fact that Cook's condition of expressiveness of \(\mathcal{L} \) is satisfied by universes with finite domains.

The research of Lipton and Snyder \([17] \) and Lipton \([16] \) culminates in a generalization and extension of Clarke's results, with a theorem (Theorem 1 in \([16] \)) which seems to tie up as equivalent the two properties of a programming language: (1) having a decidable halting problem over finite universes, and (2) the set of formulae \(P \Rightarrow [\alpha]Q \) over it being r.e. in the set of all \(U\)-valid \(L\)-wffs, for any \(U \) such that \(\mathcal{L} \) is \(U\)-expressive for \(\mathcal{L}_H \).

We conclude that relaxing the requirement and requiring that \(\mathcal{AX}_U \) be \(U\)-complete only for all arithmetical universes (i.e. playing our arithmetical-completeness game) seems a reasonable thing to do even for the restricted language of partial correctness, \(\mathcal{L}_H \).

In addition, it seems that in order for axiomatizations of much richer logics like, say, DL and DL* to be relatively complete (i.e. that they work for finite universes too), the rules that involve arithmetic (i.e. rules \((J) \) and \((P) \)) would have to be modified to deal with the finite-domain case, and would probably result in a system which is far less natural and elegant.

We are of the opinion, therefore, that the finite domains crept in because (1) the concept treated most extensively by researchers in the area was partial correctness (\([\alpha]P \) essentially), and (2) a weaker kind of expressiveness is needed to ensure the existence of an elegant relatively complete axiomatization of this particular concept on its own.

Thus we feel that it is natural and beneficial to allow the integers into ones reasoning language, in order to make possible the kind of "counting" we carry out in \(P \) and \(P^+ \).
Note that by adopting the "Hoare spirit" of structured, natural axiom systems, the remark in [23, pp. 90] "if the language is expressive it is trivial to write down a complete axiom system for partial correctness" becomes irrelevant. We are not interested in a one-rule system (as described in [23], or as shown to us in more detail by A.R. Meyer for the case of arithmetic) which has built into it essentially the full description of how to Godel-encode any wff and how to construct the equivalent formula of arithmetic. Rather, we want systems for composing our formulae step by step, using various kinds of assertions on the way. Of course the proof that these systems are complete might involve relying on the expressive power of arithmetic, and hence might call upon the use of Godel encoding, in turn making "the formulae .... be less than perspicuous" [23] (as is the case with our completeness results which at various points require finding the arithmetical equivalent to formulae). Nevertheless, we believe that the construction of these systems contributes considerably to the understanding of the concepts involved and provides the framework in which the natural and intuitive proofs one might have for one's programs can be formulated.

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9. References.


