RELATION BETWEEN THE COMPLEXITY AND THE PROBABILITY OF LARGE NUMBERS

by

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Abstract.

$H(x)$, the negative logarithm of the apriori probability $M(x)$, is Levin's variant of Kolmogorov's complexity of a natural number x . Let $a(n)$ be the minimum complexity of a number larger than n , $s(n)$ the logarithm of the apriori probability of obtaining a number larger than n . It was known that

$$s(n) \le \alpha(n) \le s(n) + H(\lfloor s(n) \rfloor) .$$

We show that the second estimate is in some sense sharp.

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Relation Between the Complexity and the Probability of Large Numbers

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Let $T(p)$ be a partial recursive function defined over binary sequences with values among the natural numbers which is <u>prefixless</u>:

(a) If p_1 is a beginning segment of p_2 and $T(p_1)$ is defined then $T(p_2) = T(p_1)$

and <u>optimal</u>:

(b) for any other prefixless p.r. function T', there is a sequence p such that $T(pq) = T'(q)$ for all q.

Let $R(p)$ denote the length of the sequence p, Levin introduced the complexity

$$H(x) = \min\{\ell(p): T(p) = x)$$

as a useful variant of Kolmogorov's complexity. See e.g. [1], also Chaitin [2], Gacs [3].

We denote by $T(p;t)$ a computable "approximation" of $T(p)$: on some Turing machine computing $T(p)$, $T(p;t)$ is $T(p)$ if $T(p)$ is computed within time t, undefined otherwise, We write

$$H(x;t) = \min\{\ell(p): T(p;t) = x)$$

$$M(x) = 2^{-H(x)}, \quad M(x;t) = 2^{-H(x;t)},$$

$$s(n) = -\log\left(\sum_{i=n}^{\infty} M(i)\right)$$

$$a(n) = \min_{i>n} H(i) .$$

$\alpha(n)$ and $s(n)$, two extremely slowly (slower than any unbounded, recursive function) growing functions, are closely related. It is known that

(1) $$s(n) \leq \alpha(n) \lesssim s(n) + H(\lfloor s(n) \rfloor) ,$$

where $\lesssim$ and $\asymp$ denote inequality and equality to within an additive, $\lesssim$ and $\approx$ to within a multiplicative constant.

The first inequality is trivial, the second one is well-known (see e.g. [4]). A hint to the proof: to find a number $\geq n$, we have only to know $2^{-s(0)}$ to within an error term $2^{-s(n)}$.

We will show that the second estimate in (1) is sharp.

Theorem. Let $g(n)$ be any positive, monotone recursive function such that

(2) $$\sum_n 2^{-g(n)} = \infty .$$

Then $a(n) > s(n) + g(s(n))$ infinitely often.

Proof. It is well-known (see e.g. [3]) that, if $\mu(n;t)$ is a computable nonnegative rational function over pairs of natural numbers, monotone in t and $\sum_n \mu(n;t) \leq 1$, i.e., for each t, $\mu(n;t)$ is a semimeasure, then

$$\mu(n;t) \lesssim M(n) .$$

Put

$$s(n;t) = \sum_{i \geq n} M(i;t)$$

$$s_\mu(n;t) = \sum_{i \geq n} \mu(i;t)$$

$$m(k;t) = \max\{n: s(n;t) < k\}$$

$$m_\mu(b;t) = \max\{n; s_\mu(n;t) < k\} .$$

The construction depends on n_k , a fast-growing recursive sequence. We will see at the end of the proof, how we should choose it in dependence of g .

Let $\mu(n;0) = 0$.

Suppose that $\mu(n;t)$ is already constructed. Put

$$k(t) = \max\{k \geq -\log(1 - s_\mu(0;t)): \exists i \in [n_{k-2}+1, n_{k-1}]$$

$$\alpha(m_\mu(i - g(i);t);t) > i\} . \qquad (3)$$

Put $n(t) = n_{k(t)}$. Let $j(t) = \max\{j: \mu(j;t) > 0\}$. Put

$$\mu(j(t)+1;t) = 2^{-n(t)}$$

$$\mu(j;t+1) = \mu(j;t) \quad \text{for } j \neq j(t) .$$

We will show that there are infinitely many i's such that for almost all t , (3) holds.

This implies, of course, that

$$\alpha(m_\mu(i - g(i)) > i .$$

That is, for some n , with

$$i-g(i) > s_\mu(n)$$

$$a(n) > i > s_\mu(n) + g(i) \geq s(n) + g(i) \geq s(n) + g(s(n))$$

and the theorem will be proved.

Suppose that, on the contrary, there is a largest i_0 among the i such that (3) holds for almost all t and a least t_0 such that (3) holds for i_0 and all $t \geq t_0$.

Under the above assumptions,

$$s_\mu(0;t) \to 1 \quad .$$

Therefore

$$\sum_t 2^{-n(t)} = 1 \quad .$$

Notation. $A(t_1,t_2) = \sum_{t=t_1}^{t_2} 2^{-n(t)}$;

$$B(t_1,t_2,k_0) = \sum \{2^{-n(t)}: t \in [t_1,t_2] , k(t) = k_0\} .$$

Lemma. There exists a triple (k_0,t_1,t_2) with $k_0 \geq k(t_0)$, $t_2 \geq t_1 \geq t_0$ such that

(a) $k(t) \geq k_0$ for $t \in [t_1,t_2]$;

(b) $2^{-n_{k_0}-1} \leq A(t_1,t_2) \leq 3 \; B(t_1,t_2,k_0)$.

Proof. For some t^0 , $(k(t_0), t_0, t^0)$ will satisfy (a) and the first inequality of (b).

Let us say that $(k_0,t_1,t_2) < (k_0',t_1',t_2')$ if $k_0' \leq k_0$, $t_1' \leq t_1 \leq t_2 \leq t_2'$.

Let (k_0,t_1,t_2) be a minimal triple $\leq (k(t_0), t_0,t^0)$, among the triples satisfying (a) and the first part of (b).

(A) For $t_3 \in [t_1,t_2]$ we have $k(t) = k_0$, otherwise the triple is not minimal.

For similar reasons we have

(B) If $t_1 \leq t_1' \leq t_2' \leq t_2$ and $k(t) > k_0$ in $[t_1',t_2']$ then then $B(t_1',t_2') < 2^{-n_{k_0}}$.

Therefore we have

$$A(t_1,t_2) \le B(t_1,t_2,k_0) + (1+\#\{t\in[t_1,t_2]: k(t) = k_0\}\cdot 2^{-n_{k_0}}$$

$$\le 2B(t_1,t_2,k_0) + 2^{-n_{k_0}} . \qquad \square$$

We concentrate now on a triple $(k,t_1,t_2) \le (k(t_0),t_0,t^0)$ satisfying (a) and (b).

Notation. For $i \in [n_{k-1},n_k]$ put

$$E_i = \{t\in[t_1,t_2]: \exists n\ H(n;t) < i, H(n;t) < H(n;t-1)\} .$$

We now estimate $s_i = \# E_i$ from below (see (5)). Let us write $E_i = \{t_{i1},t_{i2},\dots,t_{is_i}\}$, where $t_{ij} < t_{ij+1}$. Put $t_{i0} = t_1-1$, $t_{is_i+1} = t_2$. Let u_{ij} = the last t in $[t_{ij}+1,t_{ij+1}]$ (if any) with $k(t) = k$. If there is no one, $u_{ij} = t_{ij}$.

Let $\sigma_{ij} = \sum_{t=t_{ij}+1}^{u_{ij}-1} 2^{-n(t)}$, $\lambda_{ij} = -\log\sigma_{ij}$. Then by our algorithm we have

$$\alpha(m_\mu(i-g(i))\ ;u_{ij}-1) \le i .$$

On the other hand, by the definition of u_{ij},

$$\alpha(j(t_{ij}+1)\ ;u_{ij}-1) > i .$$

Therefore we have

$$\lambda_{ij} = s(j(t_{ij}+1)\ ;u_{ij}-1 \ge i-g(i) ,$$

(4) $$\sigma_{ij} \le 2^{-i+g(i)} .$$

On the other hand,

$$2^{-n_{k-1}} < \sum_{t=t_0}^{t_2} 2^{-n(t)} = \sum_{t\in E_i} 2^{-n(t)} + \sum_j \sigma_{ij} + B(t_1,t_2,k)$$

$$< s_i\cdot 2^{-n_k} + (s_i+1)2^{-i+g(i)} + B(t_1,t_2,k) .$$

Using (b) of the Lemma,

$$\frac{2}{3}\cdot 2^{-n_{k-1}} \leq (s_i+1)(2^{-n_k}+2^{-i+g(i)}) \leq 2(s_i+1)(2^{-i+g(i)}) ,$$

Hence

$$s_i \geq \frac{1}{3}\cdot 2^{-n_{k-1}+i-g(i)} - 1 ,$$

that is, for $i-g(i) > n_{k-1}+2$:

$$(5) \qquad s_i \geq \frac{1}{4}\cdot 2^{-n_{k-1}+i-g(i)} .$$

Put $m_k = \min\{i: i-g(i) > n_{k-1}+2\}$.

We have

$$1 \geq s(0;t_2) - s(0;t_1) \geq \sum_{i=m_k+1}^{n_k} \cdot 2^{-i}\cdot(s_i-s_{i-1}) + 2^{-m_k}\cdot s_{m_k}$$

$$= \sum_{i=m_k}^{n_k} \cdot 2^{-i} s_i - \sum_{i=m_k}^{n_k-1} 2^{-i-1}\cdot s_i$$

$$> \sum_{i=m_k}^{n_k-1} 2^{-i-1}\cdot s_i \geq \frac{1}{8}\cdot 2^{-n_{k-1}}\cdot \sum_{i=m_k}^{n_k} 2^{-g(i)} .$$

If n_k is chosen far enough from n_{k-1}, this will obviously lead to a contradiction. □

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